

Part I

First-order Logic

Chapter 1

Syntax and Semantics

1.1 Introduction

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In order to develop the theory and metatheory of first-order logic, we must first define the syntax and semantics of its expressions. The expressions of first-order logic are terms and **formulas**. Terms are formed from **variables**, **constant symbols**, and **function symbols**. **Formulas**, in turn, are formed from **predicate symbols** together with terms (these form the smallest, “atomic” **formulas**), and then from atomic **formulas** we can form more complex ones using logical connectives and quantifiers. There are many different ways to set down the formation rules; we give just one possible one. Other systems will chose different symbols, will select different sets of connectives as primitive, will use parentheses differently (or even not at all, as in the case of so-called Polish notation). What all approaches have in common, though, is that the formation rules define the set of terms and **formulas** *inductively*. If done properly, every expression can result essentially in only one way according to the formation rules. The inductive definition resulting in expressions that are *uniquely readable* means we can give meanings to these expressions using the same method—inductive definition.

Giving the meaning of expressions is the domain of semantics. The central concept in semantics is that of satisfaction in **a structure**. **A structure** gives meaning to the building blocks of the language: **a domain** is a non-empty set of objects. The quantifiers are interpreted as ranging over this domain, **constant symbols** are assigned elements in the domain, **function symbols** are assigned functions from the **domain** to itself, and **predicate symbols** are assigned relations on the **domain**. The **domain** together with assignments to the basic vocabulary constitutes **a structure**. **Variables** may appear in **formulas**, and in order to give a semantics, we also have to assign **elements** of the **domain** to them—this is a variable assignment. The satisfaction relation, finally, brings these together. **A formula** may be satisfied in **a structure** \mathfrak{M} relative to a variable assignment s , written as $\mathfrak{M}, s \models \varphi$. This relation is also defined by induction on the structure of φ , using the truth tables for the logical connectives to define, say, satisfaction of $\varphi \wedge \psi$ in terms of satisfaction (or not) of φ and

ψ . It then turns out that the variable assignment is irrelevant if the **formula** φ is a **sentence**, i.e., has no free variables, and so we can talk of **sentences** being simply satisfied (or not) in **structures**.

On the basis of the satisfaction relation $\mathfrak{M} \models \varphi$ for sentences we can then define the basic semantic notions of validity, entailment, and satisfiability. A sentence is valid, $\models \varphi$, if every structure satisfies it. It is entailed by a set of **sentences**, $\Gamma \models \varphi$, if every **structure** that satisfies all the **sentences** in Γ also satisfies φ . And a set of sentences is satisfiable if some **structure** satisfies all **sentences** in it at the same time. Because **formulas** are inductively defined, and satisfaction is in turn defined by induction on the structure of **formulas**, we can use induction to prove properties of our semantics and to relate the semantic notions defined.

1.2 First-Order Languages

Expressions of first-order logic are built up from a basic vocabulary containing *variables*, *constant symbols*, *predicate symbols* and sometimes *function symbols*. From them, together with logical connectives, quantifiers, and punctuation symbols such as parentheses and commas, *terms* and *formulas* are formed.

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explanation

Informally, **predicate symbols** are names for properties and relations, **constant symbols** are names for individual objects, and **function symbols** are names for mappings. These, except for the **identity predicate** $=$, are the *non-logical symbols* and together make up a language. Any first-order language \mathcal{L} is determined by its non-logical symbols. In the most general case, \mathcal{L} contains infinitely many symbols of each kind.

In the general case, we make use of the following symbols in first-order logic:

1. Logical symbols
 - a) Logical connectives: \neg (negation), \wedge (conjunction), \vee (disjunction), \rightarrow (**conditional**), \leftrightarrow (**biconditional**), \forall (universal quantifier), \exists (existential quantifier).
 - b) The propositional constant for **falsity** \perp .
 - c) The propositional constant for **truth** \top .
 - d) The two-place **identity predicate** $=$.
 - e) A **denumerable** set of **variables**: v_0, v_1, v_2, \dots
2. Non-logical symbols, making up the *standard language* of first-order logic
 - a) A **denumerable** set of n -place **predicate symbols** for each $n > 0$: $A_0^n, A_1^n, A_2^n, \dots$
 - b) A **denumerable** set of **constant symbols**: c_0, c_1, c_2, \dots
 - c) A **denumerable** set of n -place **function symbols** for each $n > 0$: $f_0^n, f_1^n, f_2^n, \dots$

3. Punctuation marks: (,), and the comma.

Most of our definitions and results will be formulated for the full standard language of first-order logic. However, depending on the application, we may also restrict the language to only a few **predicate symbols**, **constant symbols**, and **function symbols**.

Example 1.1. The language \mathcal{L}_A of arithmetic contains a single two-place **predicate symbol** $<$, a single **constant symbol** 0 , one one-place **function symbol** $!$, and two two-place **function symbols** $+$ and \times .

Example 1.2. The language of set theory \mathcal{L}_Z contains only the single two-place **predicate symbol** \in .

Example 1.3. The language of orders \mathcal{L}_{\leq} contains only the two-place **predicate symbol** \leq .

Again, these are conventions: officially, these are just aliases, e.g., $<$, \in , and \leq are aliases for A_0^2 , 0 for c_0 , $!$ for f_0^1 , $+$ for f_0^2 , \times for f_1^2 .

You may be familiar with different terminology and symbols than the ones intro we use above. Logic texts (and teachers) commonly use either \sim , \neg , and $!$ for “negation”, \wedge , \cdot , and $\&$ for “conjunction”. Commonly used symbols for the “conditional” or “implication” are \rightarrow , \Rightarrow , and \supset . Symbols for “biconditional,” “bi-implication,” or “(material) equivalence” are \leftrightarrow , \Leftrightarrow , and \equiv . The \perp symbol is variously called “falsity,” “falsum,” “absurdity,” or “bottom.” The \top symbol is variously called “truth,” “verum,” or “top.”

It is conventional to use lower case letters (e.g., a , b , c) from the beginning of the Latin alphabet for **constant symbols** (sometimes called names), and lower case letters from the end (e.g., x , y , z) for **variables**. Quantifiers combine with **variables**, e.g., x ; notational variations include $\forall x$, $(\forall x)$, (x) , Πx , \bigwedge_x for the universal quantifier and $\exists x$, $(\exists x)$, (Ex) , Σx , \bigvee_x for the existential quantifier.

We might treat all the propositional operators and both quantifiers as primitive symbols of the language. explanation We might instead choose a smaller stock of primitive symbols and treat the other **logical operators** as defined. “Truth functionally complete” sets of Boolean operators include $\{\neg, \vee\}$, $\{\neg, \wedge\}$, and $\{\neg, \rightarrow\}$ —these can be combined with either quantifier for an expressively complete first-order language.

You may be familiar with two other **logical operators**: the Sheffer stroke $|$ (named after Henry Sheffer), and Peirce’s arrow \downarrow , also known as Quine’s dagger. When given their usual readings of “nand” and “nor” (respectively), these operators are truth functionally complete by themselves.

1.3 Terms and Formulas

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sec Once a first-order language \mathcal{L} is given, we can define expressions built up from the basic vocabulary of \mathcal{L} . These include in particular *terms* and *formulas*.

Definition 1.4 (Terms). The set of *terms* $\text{Trm}(\mathcal{L})$ of \mathcal{L} is defined inductively fol:syn:frm:
defn:terms by:

1. Every **variable** is a term.
2. Every **constant symbol** of \mathcal{L} is a term.
3. If f is an n -place **function symbol** and t_1, \dots, t_n are terms, then $f(t_1, \dots, t_n)$ is a term.
4. Nothing else is a term.

A term containing no **variables** is a *closed term*.

explanation The **constant symbols** appear in our specification of the language and the terms as a separate category of symbols, but they could instead have been included as zero-place **function symbols**. We could then do without the second clause in the definition of terms. We just have to understand $f(t_1, \dots, t_n)$ as just f by itself if $n = 0$.

Definition 1.5 (Formula). The set of *formulas* $\text{Frm}(\mathcal{L})$ of the language \mathcal{L} is defined inductively as follows: fol:syn:frm:
defn:formulas

1. \perp is an atomic **formula**.
2. \top is an atomic **formula**.
3. If R is an n -place **predicate symbol** of \mathcal{L} and t_1, \dots, t_n are terms of \mathcal{L} , then $R(t_1, \dots, t_n)$ is an atomic **formula**.
4. If t_1 and t_2 are terms of \mathcal{L} , then $=(t_1, t_2)$ is an atomic **formula**.
5. If φ is a **formula**, then $\neg\varphi$ is **formula**.
6. If φ and ψ are **formulas**, then $(\varphi \wedge \psi)$ is a **formula**.
7. If φ and ψ are **formulas**, then $(\varphi \vee \psi)$ is a **formula**.
8. If φ and ψ are **formulas**, then $(\varphi \rightarrow \psi)$ is a **formula**.
9. If φ and ψ are **formulas**, then $(\varphi \leftrightarrow \psi)$ is a **formula**.
10. If φ is a **formula** and x is a **variable**, then $\forall x \varphi$ is a **formula**.
11. If φ is a **formula** and x is a **variable**, then $\exists x \varphi$ is a **formula**.
12. Nothing else is a **formula**.

The definitions of the set of terms and that of **formulas** are *inductive definitions*. Essentially, we construct the set of **formulas** in infinitely many stages. In the initial stage, we pronounce all atomic formulas to be formulas; this corresponds to the first few cases of the definition, i.e., the cases for \top , \perp , $R(t_1, \dots, t_n)$ and $=(t_1, t_2)$. “Atomic **formula**” thus means any **formula** of this form. explanation

The other cases of the definition give rules for constructing new **formulas** out of **formulas** already constructed. At the second stage, we can use them to construct **formulas** out of atomic **formulas**. At the third stage, we construct new formulas from the atomic formulas and those obtained in the second stage, and so on. A **formula** is anything that is eventually constructed at such a stage, and nothing else.

By convention, we write $=$ between its arguments and leave out the parentheses: $t_1 = t_2$ is an abbreviation for $=(t_1, t_2)$. Moreover, $\neg=(t_1, t_2)$ is abbreviated as $t_1 \neq t_2$. When writing a formula $(\psi * \chi)$ constructed from ψ , χ using a two-place connective $*$, we will often leave out the outermost pair of parentheses and write simply $\psi * \chi$.

Some logic texts require that the **variable** x must occur in φ in order for $\exists x\varphi$ and $\forall x\varphi$ to count as **formulas**. Nothing bad happens if you don’t require this, and it makes things easier. intro

If we work in a language for a specific application, we will often write two-place **predicate symbols** and **function symbols** between the respective terms, e.g., $t_1 < t_2$ and $(t_1 + t_2)$ in the language of arithmetic and $t_1 \in t_2$ in the language of set theory. The successor function in the language of arithmetic is even written conventionally *after* its argument: t' . Officially, however, these are just conventional abbreviations for $A_0^2(t_1, t_2)$, $f_0^2(t_1, t_2)$, $A_0^2(t_1, t_2)$ and $f_0^1(t)$, respectively.

Definition 1.6 (Syntactic identity). The symbol \equiv expresses syntactic identity between strings of symbols, i.e., $\varphi \equiv \psi$ iff φ and ψ are strings of symbols of the same length and which contain the same symbol in each place.

The \equiv symbol may be flanked by strings obtained by concatenation, e.g., $\varphi \equiv (\psi \vee \chi)$ means: the string of symbols φ is the same string as the one obtained by concatenating an opening parenthesis, the string ψ , the \vee symbol, the string χ , and a closing parenthesis, in this order. If this is the case, then we know that the first symbol of φ is an opening parenthesis, φ contains ψ as a substring (starting at the second symbol), that substring is followed by \vee , etc.

1.4 Unique Readability

The way we defined **formulas** guarantees that every **formula** has a *unique reading*, i.e., there is essentially only one way of constructing it according to our formation rules for **formulas** and only one way of “interpreting” it. If this were not so, we would have ambiguous **formulas**, i.e., **formulas** that have more than one reading or interpretation—and that is clearly something we want to explanation

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sec

avoid. But more importantly, without this property, most of the definitions and proofs we are going to give will not go through.

Perhaps the best way to make this clear is to see what would happen if we had given bad rules for forming **formulas** that would not guarantee unique readability. For instance, we could have forgotten the parentheses in the formation rules for connectives, e.g., we might have allowed this:

If φ and ψ are **formulas**, then so is $\varphi \rightarrow \psi$.

Starting from an atomic formula θ , this would allow us to form $\theta \rightarrow \theta$. From this, together with θ , we would get $\theta \rightarrow \theta \rightarrow \theta$. But there are two ways to do this:

1. We take θ to be φ and $\theta \rightarrow \theta$ to be ψ .
2. We take φ to be $\theta \rightarrow \theta$ and ψ is θ .

Correspondingly, there are two ways to “read” the **formula** $\theta \rightarrow \theta \rightarrow \theta$. It is of the form $\psi \rightarrow \chi$ where ψ is θ and χ is $\theta \rightarrow \theta$, but *it is also* of the form $\psi \rightarrow \chi$ with ψ being $\theta \rightarrow \theta$ and χ being θ .

If this happens, our definitions will not always work. For instance, when we define the **main operator** of a formula, we say: in a formula of the form $\psi \rightarrow \chi$, the **main operator** is the indicated occurrence of \rightarrow . But if we can match the formula $\theta \rightarrow \theta \rightarrow \theta$ with $\psi \rightarrow \chi$ in the two different ways mentioned above, then in one case we get the first occurrence of \rightarrow as the **main operator**, and in the second case the second occurrence. But we intend the **main operator** to be a *function* of the **formula**, i.e., every **formula** must have exactly one **main operator** occurrence.

Lemma 1.7. *The number of left and right parentheses in a **formula** φ are equal.*

Proof. We prove this by induction on the way φ is constructed. This requires two things: (a) We have to prove first that all atomic formulas have the property in question (the induction basis). (b) Then we have to prove that when we construct new formulas out of given formulas, the new formulas have the property provided the old ones do.

Let $l(\varphi)$ be the number of left parentheses, and $r(\varphi)$ the number of right parentheses in φ , and $l(t)$ and $r(t)$ similarly the number of left and right parentheses in a term t . We leave the proof that for any term t , $l(t) = r(t)$ as an exercise.

1. $\varphi \equiv \perp$: φ has 0 left and 0 right parentheses.
2. $\varphi \equiv \top$: φ has 0 left and 0 right parentheses.
3. $\varphi \equiv R(t_1, \dots, t_n)$: $l(\varphi) = 1 + l(t_1) + \dots + l(t_n) = 1 + r(t_1) + \dots + r(t_n) = r(\varphi)$. Here we make use of the fact, left as an exercise, that $l(t) = r(t)$ for any term t .

4. $\varphi \equiv t_1 = t_2$: $l(\varphi) = l(t_1) + l(t_2) = r(t_1) + r(t_2) = r(\varphi)$.
5. $\varphi \equiv \neg\psi$: By induction hypothesis, $l(\psi) = r(\psi)$. Thus $l(\varphi) = l(\psi) = r(\psi) = r(\varphi)$.
6. $\varphi \equiv (\psi * \chi)$: By induction hypothesis, $l(\psi) = r(\psi)$ and $l(\chi) = r(\chi)$. Thus $l(\varphi) = 1 + l(\psi) + l(\chi) = 1 + r(\psi) + r(\chi) = r(\varphi)$.
7. $\varphi \equiv \forall x \psi$: By induction hypothesis, $l(\psi) = r(\psi)$. Thus, $l(\varphi) = l(\psi) = r(\psi) = r(\varphi)$.
8. $\varphi \equiv \exists x \psi$: Similarly.

□

Definition 1.8 (Proper prefix). A string of symbols ψ is a *proper prefix* of a string of symbols φ if concatenating ψ and a non-empty string of symbols yields φ .

fol:syn:unq;
lem:no-prefix **Lemma 1.9.** *If φ is a formula, and ψ is a proper prefix of φ , then ψ is not a formula.*

Proof. Exercise.

□

Problem 1.1. Prove Lemma 1.9.

fol:syn:unq;
prop:unique-atomic **Proposition 1.10.** *If φ is an atomic formula, then it satisfies one, and only one of the following conditions.*

1. $\varphi \equiv \perp$.
2. $\varphi \equiv \top$.
3. $\varphi \equiv R(t_1, \dots, t_n)$ where R is an n -place predicate symbol, t_1, \dots, t_n are terms, and each of R, t_1, \dots, t_n is uniquely determined.
4. $\varphi \equiv t_1 = t_2$ where t_1 and t_2 are uniquely determined terms.

Proof. Exercise.

□

Problem 1.2. Prove Proposition 1.10 (Hint: Formulate and prove a version of Lemma 1.9 for terms.)

Proposition 1.11 (Unique Readability). *Every formula satisfies one, and only one of the following conditions.*

1. φ is atomic.
2. φ is of the form $\neg\psi$.
3. φ is of the form $(\psi \wedge \chi)$.
4. φ is of the form $(\psi \vee \chi)$.

5. φ is of the form $(\psi \rightarrow \chi)$.
6. φ is of the form $(\psi \leftrightarrow \chi)$.
7. φ is of the form $\forall x \psi$.
8. φ is of the form $\exists x \psi$.

Moreover, in each case ψ , or ψ and χ , are uniquely determined. This means that, e.g., there are no different pairs ψ, χ and ψ', χ' so that φ is both of the form $(\psi \rightarrow \chi)$ and $(\psi' \rightarrow \chi')$.

Proof. The formation rules require that if a **formula** is not atomic, it must start with an opening parenthesis ($($, \neg , or with a quantifier. On the other hand, every **formula** that start with one of the following symbols must be atomic: a **predicate symbol**, a **function symbol**, a **constant symbol**, \perp , \top .

So we really only have to show that if φ is of the form $(\psi * \chi)$ and also of the form $(\psi' *' \chi')$, then $\psi \equiv \psi'$, $\chi \equiv \chi'$, and $* = *'$.

So suppose both $\varphi \equiv (\psi * \chi)$ and $\varphi \equiv (\psi' *' \chi')$. Then either $\psi \equiv \psi'$ or not. If it is, clearly $* = *'$ and $\chi \equiv \chi'$, since they then are substrings of φ that begin in the same place and are of the same length. The other case is $\psi \not\equiv \psi'$. Since ψ and ψ' are both substrings of φ that begin at the same place, one must be a proper prefix of the other. But this is impossible by [Lemma 1.9](#). \square

1.5 Main operator of a Formula

explanation

It is often useful to talk about the last operator used in constructing a **formula** φ . This operator is called the *main operator* of φ . Intuitively, it is the “outermost” operator of φ . For example, the main operator of $\neg\varphi$ is \neg , the main operator of $(\varphi \vee \psi)$ is \vee , etc.

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Definition 1.12 (Main operator). The *main operator* of a **formula** φ is defined as follows:

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def:main-op

1. φ is atomic: φ has no **main operator**.
2. $\varphi \equiv \neg\psi$: the **main operator** of φ is \neg .
3. $\varphi \equiv (\psi \wedge \chi)$: the **main operator** of φ is \wedge .
4. $\varphi \equiv (\psi \vee \chi)$: the **main operator** of φ is \vee .
5. $\varphi \equiv (\psi \rightarrow \chi)$: the **main operator** of φ is \rightarrow .
6. $\varphi \equiv (\psi \leftrightarrow \chi)$: the **main operator** of φ is \leftrightarrow .
7. $\varphi \equiv \forall x \psi$: the **main operator** of φ is \forall .
8. $\varphi \equiv \exists x \psi$: the **main operator** of φ is \exists .

In each case, we intend the specific indicated *occurrence* of the **main operator** in the formula. For instance, since the formula $((\theta \rightarrow \alpha) \rightarrow (\alpha \rightarrow \theta))$ is of the form $(\psi \rightarrow \chi)$ where ψ is $(\theta \rightarrow \alpha)$ and χ is $(\alpha \rightarrow \theta)$, the second occurrence of \rightarrow is the **main operator**.

This is a *recursive* definition of a function which maps all non-atomic **formulas** to their **main operator** occurrence. Because of the way **formulas** are defined inductively, every **formula** φ satisfies one of the cases in [Definition 1.12](#). This guarantees that for each non-atomic **formula** φ a **main operator** exists. Because each **formula** satisfies only one of these conditions, and because the smaller **formulas** from which φ is constructed are uniquely determined in each case, the **main operator** occurrence of φ is unique, and so we have defined a function. explanation

We call **formulas** by the following names depending on which symbol their **main operator** is:

Main operator	Type of formula	Example
none	atomic (formula)	$\perp, \top, R(t_1, \dots, t_n), t_1 = t_2$
\neg	negation	$\neg\varphi$
\wedge	conjunction	$(\varphi \wedge \psi)$
\vee	disjunction	$(\varphi \vee \psi)$
\rightarrow	conditional	$(\varphi \rightarrow \psi)$
\forall	universal (formula)	$\forall x \varphi$
\exists	existential (formula)	$\exists x \varphi$

1.6 Subformulas

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sec

It is often useful to talk about the **formulas** that “make up” a given **formula**. We call these its *subformulas*. Any **formula** counts as a **subformula** of itself; a **subformula** of φ other than φ itself is a *proper subformula*. explanation

Definition 1.13 (Immediate **Subformula**). If φ is a **formula**, the *immediate subformulas* of φ are defined inductively as follows:

- Atomic **formulas** have no immediate **subformulas**.
- $\varphi \equiv \neg\psi$: The only immediate **subformula** of φ is ψ .
- $\varphi \equiv (\psi * \chi)$: The immediate **subformulas** of φ are ψ and χ ($*$ is any one of the two-place connectives).
- $\varphi \equiv \forall x \psi$: The only immediate **subformula** of φ is ψ .
- $\varphi \equiv \exists x \psi$: The only immediate **subformula** of φ is ψ .

Definition 1.14 (Proper **Subformula**). If φ is a **formula**, the *proper subformulas* of φ are recursively as follows:

- Atomic **formulas** have no proper **subformulas**.
- $\varphi \equiv \neg\psi$: The proper **subformulas** of φ are ψ together with all proper **subformulas** of ψ .

3. $\varphi \equiv (\psi * \chi)$: The proper **subformulas** of φ are ψ , χ , together with all proper **subformulas** of ψ and those of χ .
4. $\varphi \equiv \forall x \psi$: The proper **subformulas** of φ are ψ together with all proper **subformulas** of ψ .
5. $\varphi \equiv \exists x \psi$: The proper **subformulas** of φ are ψ together with all proper **subformulas** of ψ .

Definition 1.15 (Subformula). The **subformulas** of φ are φ itself together with all its proper **subformulas**.

explanation

Note the subtle difference in how we have defined immediate **subformulas** and proper **subformulas**. In the first case, we have directly defined the immediate **subformulas** of a formula φ for each possible form of φ . It is an explicit definition by cases, and the cases mirror the inductive definition of the set of **formulas**. In the second case, we have also mirrored the way the set of all **formulas** is defined, but in each case we have also included the proper **subformulas** of the smaller **formulas** ψ , χ in addition to these **formulas** themselves. This makes the definition *recursive*. In general, a definition of a function on an inductively defined set (in our case, **formulas**) is recursive if the cases in the definition of the function make use of the function itself. To be well defined, we must make sure, however, that we only ever use the values of the function for arguments that come “before” the one we are defining—in our case, when defining “proper **subformula**” for $(\psi * \chi)$ we only use the proper **subformulas** of the “earlier” **formulas** ψ and χ .

1.7 Free Variables and Sentences

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Definition 1.16 (Free occurrences of a variable). The *free* occurrences of a **variable** in a **formula** are defined inductively as follows:

fol:syn:fvs:
defn:free-occ

1. φ is atomic: all **variable** occurrences in φ are free.
2. $\varphi \equiv \neg\psi$: the free **variable** occurrences of φ are exactly those of ψ .
3. $\varphi \equiv (\psi * \chi)$: the free **variable** occurrences of φ are those in ψ together with those in χ .
4. $\varphi \equiv \forall x \psi$: the free **variable** occurrences in φ are all of those in ψ except for occurrences of x .
5. $\varphi \equiv \exists x \psi$: the free **variable** occurrences in φ are all of those in ψ except for occurrences of x .

Definition 1.17 (Bound Variables). An occurrence of a **variable** in a formula φ is *bound* if it is not free.

Problem 1.3. Give an inductive definition of the bound variable occurrences along the lines of [Definition 1.16](#).

Definition 1.18 (Scope). If $\forall x \psi$ is an occurrence of a subformula in a formula φ , then the corresponding occurrence of ψ in φ is called the *scope* of the corresponding occurrence of $\forall x$. Similarly for $\exists x$.

If ψ is the scope of a quantifier occurrence $\forall x$ or $\exists x$ in φ , then all occurrences of x which are free in ψ are said to be *bound by* the mentioned quantifier occurrence.

Example 1.19. Consider the following formula:

$$\exists v_0 \underbrace{A_0^2(v_0, v_1)}_{\psi}$$

ψ represents the scope of $\exists v_0$. The quantifier binds the occurrence of v_0 in ψ , but does not bind the occurrence of v_1 . So v_1 is a free variable in this case.

We can now see how this might work in a more complicated [formula](#) φ :

$$\forall v_0 \underbrace{(A_0^1(v_0) \rightarrow A_0^2(v_0, v_1))}_{\psi} \rightarrow \exists v_1 \underbrace{(A_1^2(v_0, v_1) \vee \forall v_0 \overbrace{\neg A_1^1(v_0)}^{\theta})}_{\chi}$$

ψ is the scope of the first $\forall v_0$, χ is the scope of $\exists v_1$, and θ is the scope of the second $\forall v_0$. The first $\forall v_0$ binds the occurrences of v_0 in ψ , $\exists v_1$ the occurrence of v_1 in χ , and the second $\forall v_0$ binds the occurrence of v_0 in θ . The first occurrence of v_1 and the fourth occurrence of v_0 are free in φ . The last occurrence of v_0 is free in θ , but bound in χ and φ .

Definition 1.20 (Sentence). A [formula](#) φ is a *sentence* iff it contains no free occurrences of [variables](#).

1.8 Substitution

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Definition 1.21 (Substitution in a term). We define $s[t/x]$, the result of *substituting* t for every occurrence of x in s , recursively:

1. $s \equiv c$: $s[t/x]$ is just s .
2. $s \equiv y$: $s[t/x]$ is also just s , provided y is a variable and $y \neq x$.
3. $s \equiv x$: $s[t/x]$ is t .
4. $s \equiv f(t_1, \dots, t_n)$: $s[t/x]$ is $f(t_1[t/x], \dots, t_n[t/x])$.

Definition 1.22. A term t is *free for* x in φ if none of the free occurrences of x in φ occur in the scope of a quantifier that binds a variable in t .

Example 1.23.

1. v_8 is free for v_1 in $\exists v_3 A_4^2(v_3, v_1)$
2. $f_1^2(v_1, v_2)$ is *not* free for v_0 in $\forall v_2 A_4^2(v_0, v_2)$

Definition 1.24 (Substitution in a formula). If φ is a formula, x is a variable, and t is a term free for x in φ , then $\varphi[t/x]$ is the result of substituting t for all free occurrences of x in φ .

1. $\varphi \equiv \perp$: $\varphi[t/x]$ is \perp .
2. $\varphi \equiv \top$: $\varphi[t/x]$ is \top .
3. $\varphi \equiv P(t_1, \dots, t_n)$: $\varphi[t/x]$ is $P(t_1[t/x], \dots, t_n[t/x])$.
4. $\varphi \equiv t_1 = t_2$: $\varphi[t/x]$ is $t_1[t/x] = t_2[t/x]$.
5. $\varphi \equiv \neg\psi$: $\varphi[t/x]$ is $\neg\psi[t/x]$.
6. $\varphi \equiv (\psi \wedge \chi)$: $\varphi[t/x]$ is $(\psi[t/x] \wedge \chi[t/x])$.
7. $\varphi \equiv (\psi \vee \chi)$: $\varphi[t/x]$ is $(\psi[t/x] \vee \chi[t/x])$.
8. $\varphi \equiv (\psi \rightarrow \chi)$: $\varphi[t/x]$ is $(\psi[t/x] \rightarrow \chi[t/x])$.
9. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\varphi[t/x]$ is $(\psi[t/x] \leftrightarrow \chi[t/x])$.
10. $\varphi \equiv \forall y \psi$: $\varphi[t/x]$ is $\forall y \psi[t/x]$, provided y is a variable other than x ; otherwise $\varphi[t/x]$ is just φ .
11. $\varphi \equiv \exists y \psi$: $\varphi[t/x]$ is $\exists y \psi[t/x]$, provided y is a variable other than x ; otherwise $\varphi[t/x]$ is just φ .

explanation

Note that substitution may be vacuous: If x does not occur in φ at all, then $\varphi[t/x]$ is just φ .

The restriction that t must be free for x in φ is necessary to exclude cases like the following. If $\varphi \equiv \exists y x < y$ and $t \equiv y$, then $\varphi[t/x]$ would be $\exists y y < y$. In this case the free variable y is “captured” by the quantifier $\exists y$ upon substitution, and that is undesirable. For instance, we would like it to be the case that whenever $\forall x \psi$ holds, so does $\psi[t/x]$. But consider $\forall x \exists y x < y$ (here ψ is $\exists y x < y$). It is sentence that is true about, e.g., the natural numbers: for every number x there is a number y greater than it. If we allowed y as a possible substitution for x , we would end up with $\psi[y/x] \equiv \exists y y < y$, which is false. We prevent this by requiring that none of the free variables in t would end up being bound by a quantifier in φ .

We often use the following convention to avoid cumbersome notation: If φ is a formula with a free variable x , we write $\varphi(x)$ to indicate this. When it is clear which φ and x we have in mind, and t is a term (assumed to be free for x in $\varphi(x)$), then we write $\varphi(t)$ as short for $\varphi(x)[t/x]$.

1.9 Structures for First-order Languages

fol:syn:str:
sec

First-order languages are, by themselves, *uninterpreted*: the **constant symbols**, **function symbols**, and **predicate symbols** have no specific meaning attached to them. Meanings are given by specifying a *structure*. It specifies the *domain*, i.e., the objects which the **constant symbols** pick out, the **function symbols** operate on, and the quantifiers range over. In addition, it specifies which **constant symbols** pick out which objects, how a **function symbol** maps objects to objects, and which objects the **predicate symbols** apply to. **Structures** are the basis for *semantic* notions in logic, e.g., the notion of consequence, validity, satisfiability. They are variously called “structures,” “interpretations,” or “models” in the literature.

explanation

Definition 1.25 (Structures). A *structure* \mathfrak{M} , for a language \mathcal{L} of first-order logic consists of the following elements:

1. *Domain*: a non-empty set, $|\mathfrak{M}|$
2. *Interpretation of constant symbols*: for each **constant symbol** c of \mathcal{L} , an **element** $c^{\mathfrak{M}} \in |\mathfrak{M}|$
3. *Interpretation of predicate symbols*: for each n -place **predicate symbol** R of \mathcal{L} (other than $=$), an n -place relation $R^{\mathfrak{M}} \subseteq |\mathfrak{M}|^n$
4. *Interpretation of function symbols*: for each n -place **function symbol** f of \mathcal{L} , an n -place function $f^{\mathfrak{M}}: |\mathfrak{M}|^n \rightarrow |\mathfrak{M}|$

Example 1.26. A *structure* \mathfrak{N} for the language of arithmetic consists of a set, an element of $|\mathfrak{N}|$, $o^{\mathfrak{N}}$, as interpretation of the **constant symbol** o , a one-place function $\iota^{\mathfrak{N}}: |\mathfrak{N}| \rightarrow |\mathfrak{N}|$, two two-place functions $+^{\mathfrak{N}}$ and $\times^{\mathfrak{N}}$, both $|\mathfrak{N}|^2 \rightarrow |\mathfrak{N}|$, and a two-place relation $<^{\mathfrak{N}} \subseteq |\mathfrak{N}|^2$.

An obvious example of such a structure is the following:

1. $|\mathfrak{N}| = \mathbb{N}$
2. $o^{\mathfrak{N}} = 0$
3. $\iota^{\mathfrak{N}}(n) = n + 1$ for all $n \in \mathbb{N}$
4. $+^{\mathfrak{N}}(n, m) = n + m$ for all $n, m \in \mathbb{N}$
5. $\times^{\mathfrak{N}}(n, m) = n \cdot m$ for all $n, m \in \mathbb{N}$
6. $<^{\mathfrak{N}} = \{(n, m) : n \in \mathbb{N}, m \in \mathbb{N}, n < m\}$

The structure \mathfrak{N} for \mathcal{L}_A so defined is called the *standard model of arithmetic*, because it interprets the non-logical constants of \mathcal{L}_A exactly how you would expect.

However, there are many other possible **structures** for \mathcal{L}_A . For instance, we might take as the domain the set \mathbb{Z} of integers instead of \mathbb{N} , and define the interpretations of o , ι , $+$, \times , $<$ accordingly. But we can also define structures for \mathcal{L}_A which have nothing even remotely to do with numbers.

Example 1.27. A structure \mathfrak{M} for the language \mathcal{L}_Z of set theory requires just a set and a single-two place relation. So technically, e.g., the set of people plus the relation “ x is older than y ” could be used as a **structure** for \mathcal{L}_Z , as well as \mathbb{N} together with $n \geq m$ for $n, m \in \mathbb{N}$.

A particularly interesting **structure** for \mathcal{L}_Z in which the **elements** of the domain are actually sets, and the interpretation of \in actually is the relation “ x is an **element** of y ” is the **structure** $\mathfrak{H}\mathfrak{F}$ of *hereditarily finite sets*:

1. $|\mathfrak{H}\mathfrak{F}| = \emptyset \cup \wp(\emptyset) \cup \wp(\wp(\emptyset)) \cup \wp(\wp(\wp(\emptyset))) \cup \dots$;
2. $\in^{\mathfrak{H}\mathfrak{F}} = \{\langle x, y \rangle : x, y \in |\mathfrak{H}\mathfrak{F}|, x \in y\}$.

digression

The stipulations we make as to what counts as a **structure** impact our logic. For example, the choice to prevent empty domains ensures, given the usual account of satisfaction (or truth) for quantified sentences, that $\exists x (\varphi(x) \vee \neg \varphi(x))$ is valid—that is, a logical truth. And the stipulation that all **constant symbols** must refer to an object in the domain ensures that the existential generalization is a sound pattern of inference: $\varphi(a)$, therefore $\exists x \varphi(x)$. If we allowed names to refer outside the domain, or to not refer, then we would be on our way to a *free logic*, in which existential generalization requires an additional premise: $\varphi(a)$ and $\exists x x = a$, therefore $\exists x \varphi(x)$.

1.10 Covered **Structures** for First-order Languages

explanation

Recall that a term is *closed* if it contains no **variables**.

fol:syn:cov:
sec

Definition 1.28 (**Value** of closed terms). If t is a closed term of the language \mathcal{L} and \mathfrak{M} is a **structure** for \mathcal{L} , the **value** $\text{Val}^{\mathfrak{M}}(t)$ is defined as follows:

1. If t is just the **constant symbol** c , then $\text{Val}^{\mathfrak{M}}(c) = c^{\mathfrak{M}}$.
2. If t is of the form $f(t_1, \dots, t_n)$, then

$$\text{Val}^{\mathfrak{M}}(t) = f^{\mathfrak{M}}(\text{Val}^{\mathfrak{M}}(t_1), \dots, \text{Val}^{\mathfrak{M}}(t_n)).$$

Definition 1.29 (**Covered structure**). A **structure** is *covered* if every element of the domain is the **value** of some closed term.

Example 1.30. Let \mathcal{L} be the language with **constant symbols** *zero*, *one*, *two*, \dots , the binary **predicate symbol** $<$, and the binary **function symbols** $+$ and \times . Then a **structure** \mathfrak{M} for \mathcal{L} is the one with domain $|\mathfrak{M}| = \{0, 1, 2, \dots\}$ and assignments $zero^{\mathfrak{M}} = 0$, $one^{\mathfrak{M}} = 1$, $two^{\mathfrak{M}} = 2$, and so forth. For the binary relation symbol $<$, the set $<^{\mathfrak{M}}$ is the set of all pairs $\langle c_1, c_2 \rangle \in |\mathfrak{M}|^2$ such that c_1 is less than c_2 : for example, $\langle 1, 3 \rangle \in <^{\mathfrak{M}}$ but $\langle 2, 2 \rangle \notin <^{\mathfrak{M}}$. For the binary **function symbol** $+$, define $+^{\mathfrak{M}}$ in the usual way—for example, $+^{\mathfrak{M}}(2, 3)$ maps to 5, and similarly for the binary **function symbol** \times . Hence, the **value** of

four is just 4, and the **value** of $\times(\textit{two}, +(\textit{three}, \textit{zero}))$ (or in infix notation, $\textit{two} \times (\textit{three} + \textit{zero})$) is

$$\begin{aligned}
 \text{Val}^{\mathfrak{M}}(\times(\textit{two}, +(\textit{three}, \textit{zero}))) &= \\
 &= \times^{\mathfrak{M}}(\text{Val}^{\mathfrak{M}}(\textit{two}), \text{Val}^{\mathfrak{M}}(\textit{two}, +(\textit{three}, \textit{zero}))) \\
 &= \times^{\mathfrak{M}}(\text{Val}^{\mathfrak{M}}(\textit{two}), +^{\mathfrak{M}}(\text{Val}^{\mathfrak{M}}(\textit{three}), \text{Val}^{\mathfrak{M}}(\textit{zero}))) \\
 &= \times^{\mathfrak{M}}(\textit{two}^{\mathfrak{M}}, +^{\mathfrak{M}}(\textit{three}^{\mathfrak{M}}, \textit{zero}^{\mathfrak{M}})) \\
 &= \times^{\mathfrak{M}}(2, +^{\mathfrak{M}}(3, 0)) \\
 &= \times^{\mathfrak{M}}(2, 3) \\
 &= 6
 \end{aligned}$$

Problem 1.4. Is \mathfrak{N} , the standard model of arithmetic, covered? Explain.

1.11 Satisfaction of a Formula in a Structure

fol:syn:sat:
sec

The basic notion that relates expressions such as terms and **formulas**, on the one hand, and **structures** on the other, are those of **value** of a term and **satisfaction** of a **formula**. Informally, the **value** of a term is an **element** of a **structure**—if the term is just a constant, its **value** is the object assigned to the constant by the **structure**, and if it is built up using **function symbols**, the **value** is computed from the **values** of constants and the functions assigned to the functions in the term. A **formula** is **satisfied** in a **structure** if the interpretation given to the predicates makes the **formula** true in the domain of the **structure**. This notion of satisfaction is specified inductively: the specification of the **structure** directly states when atomic **formulas** are satisfied, and we define when a complex **formula** is satisfied depending on the main connective or quantifier and whether or not the immediate **subformulas** are satisfied. The case of the quantifiers here is a bit tricky, as the immediate **subformula** of a quantified **formula** has a free **variable**, and **structures** don't specify the **values** of **variables**. In order to deal with this difficulty, we also introduce *variable assignments* and define satisfaction not with respect to a **structure** alone, but with respect to a **structure** plus a **variable** assignment.

explanation

Definition 1.31 (Variable Assignment). A *variable assignment* s for a **structure** \mathfrak{M} is a function which maps each **variable** to an element of $|\mathfrak{M}|$, i.e., $s: \text{Var} \rightarrow |\mathfrak{M}|$.

A **structure** assigns a **value** to each **constant symbol**, and a variable assignment to each variable. But we want to use terms built up from them to also name **elements** of the **domain**. For this we define the **value** of terms inductively. For **constant symbols** and variables the value is just as the **structure** or the variable assignment specifies it; for more complex terms it is computed recursively using the functions the **structure** assigns to the **function symbols**.

explanation

Definition 1.32 (Value of Terms). If t is a term of the language \mathcal{L} , \mathfrak{M} is a structure for \mathcal{L} , and s is a variable assignment for \mathfrak{M} , the value $\text{Val}_s^{\mathfrak{M}}(t)$ is defined as follows:

1. $t \equiv c$: $\text{Val}_s^{\mathfrak{M}}(t) = c^{\mathfrak{M}}$.
2. $t \equiv x$: $\text{Val}_s^{\mathfrak{M}}(t) = s(x)$.
3. $t \equiv f(t_1, \dots, t_n)$:

$$\text{Val}_s^{\mathfrak{M}}(t) = f^{\mathfrak{M}}(\text{Val}_s^{\mathfrak{M}}(t_1), \dots, \text{Val}_s^{\mathfrak{M}}(t_n)).$$

Definition 1.33 (x -Variant). If s is a variable assignment for a structure \mathfrak{M} , then any variable assignment s' for \mathfrak{M} which differs from s at most in what it assigns to x is called an x -variant of s . If s' is an x -variant of s we write $s \sim_x s'$.

explanation

Note that an x -variant of an assignment s does not *have* to assign something different to x . In fact, every assignment counts as an x -variant of itself.

Definition 1.34 (Satisfaction). Satisfaction of a formula φ in a structure \mathfrak{M} relative to a variable assignment s , in symbols: $\mathfrak{M}, s \models \varphi$, is defined recursively as follows. (We write $\mathfrak{M}, s \not\models \varphi$ to mean “not $\mathfrak{M}, s \models \varphi$.”)

fol:syn:sat:
defn:satisfaction

1. $\varphi \equiv \perp$: $\mathfrak{M}, s \not\models \varphi$.
2. $\varphi \equiv \top$: $\mathfrak{M}, s \models \varphi$.
3. $\varphi \equiv R(t_1, \dots, t_n)$: $\mathfrak{M}, s \models \varphi$ iff $\langle \text{Val}_s^{\mathfrak{M}}(t_1), \dots, \text{Val}_s^{\mathfrak{M}}(t_n) \rangle \in R^{\mathfrak{M}}$.
4. $\varphi \equiv t_1 = t_2$: $\mathfrak{M}, s \models \varphi$ iff $\text{Val}_s^{\mathfrak{M}}(t_1) = \text{Val}_s^{\mathfrak{M}}(t_2)$.
5. $\varphi \equiv \neg\psi$: $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s \not\models \psi$.
6. $\varphi \equiv (\psi \wedge \chi)$: $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s \models \psi$ and $\mathfrak{M}, s \models \chi$.
7. $\varphi \equiv (\psi \vee \chi)$: $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s \models \psi$ or $\mathfrak{M}, s \models \chi$ (or both).
8. $\varphi \equiv (\psi \rightarrow \chi)$: $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s \not\models \psi$ or $\mathfrak{M}, s \models \chi$ (or both).
9. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\mathfrak{M}, s \models \varphi$ iff either both $\mathfrak{M}, s \models \psi$ and $\mathfrak{M}, s \models \chi$, or neither $\mathfrak{M}, s \models \psi$ nor $\mathfrak{M}, s \models \chi$.
10. $\varphi \equiv \forall x \psi$: $\mathfrak{M}, s \models \varphi$ iff for every x -variant s' of s , $\mathfrak{M}, s' \models \psi$.
11. $\varphi \equiv \exists x \psi$: $\mathfrak{M}, s \models \varphi$ iff there is an x -variant s' of s so that $\mathfrak{M}, s' \models \psi$.

explanation

The variable assignments are important in the last two clauses. We cannot define satisfaction of $\forall x \psi(x)$ by “for all $a \in |\mathfrak{M}|$, $\mathfrak{M} \models \psi(a)$.” We cannot define satisfaction of $\exists x \psi(x)$ by “for at least one $a \in |\mathfrak{M}|$, $\mathfrak{M} \models \psi(a)$.” The reason is that a is not symbol of the language, and so $\psi(a)$ is not a formula (that is, $\psi[a/x]$ is undefined). We also cannot assume that we have constant symbols

or terms available that name every **element** of \mathfrak{M} , since there is nothing in the definition of **structures** that requires it. Even in the standard language the set of **constant symbols** is **denumerable**, so if $|\mathfrak{M}|$ is not **enumerable** there aren't even enough **constant symbols** to name every object.

Example 1.35. Let $=\{a, b, f, R\}$ where a and b are **constant symbols**, f is a two-place **function symbol**, and R is a two-place **predicate symbol**. Consider the **structure** \mathfrak{M} defined by:

1. $|\mathfrak{M}| = \{1, 2, 3, 4\}$
2. $a^{\mathfrak{M}} = 1$
3. $b^{\mathfrak{M}} = 2$
4. $f^{\mathfrak{M}}(x, y) = x + y$ if $x + y \leq 3$ and $= 3$ otherwise.
5. $R^{\mathfrak{M}} = \{\langle 1, 1 \rangle, \langle 1, 2 \rangle, \langle 2, 3 \rangle, \langle 2, 4 \rangle\}$

The function $s(x) = 1$ that assigns $1 \in |\mathfrak{M}|$ to every **variable** is a variable assignment for \mathfrak{M} .

Then

$$\text{Val}_s^{\mathfrak{M}}(f(a, b)) = f^{\mathfrak{M}}(\text{Val}_s^{\mathfrak{M}}(a), \text{Val}_s^{\mathfrak{M}}(b)).$$

Since a and b are **constant symbols**, $\text{Val}_s^{\mathfrak{M}}(a) = a^{\mathfrak{M}} = 1$ and $\text{Val}_s^{\mathfrak{M}}(b) = b^{\mathfrak{M}} = 2$. So

$$\text{Val}_s^{\mathfrak{M}}(f(a, b)) = f^{\mathfrak{M}}(1, 2) = 1 + 2 = 3.$$

To compute the value of $f(f(a, b), a)$ we have to consider

$$\text{Val}_s^{\mathfrak{M}}(f(f(a, b), a)) = f^{\mathfrak{M}}(\text{Val}_s^{\mathfrak{M}}(f(a, b)), \text{Val}_s^{\mathfrak{M}}(a)) = f^{\mathfrak{M}}(3, 1) = 3,$$

since $3 + 1 > 3$. Since $s(x) = 1$ and $\text{Val}_s^{\mathfrak{M}}(x) = s(x)$, we also have

$$\text{Val}_s^{\mathfrak{M}}(f(f(a, b), x)) = f^{\mathfrak{M}}(\text{Val}_s^{\mathfrak{M}}(f(a, b)), \text{Val}_s^{\mathfrak{M}}(x)) = f^{\mathfrak{M}}(3, 1) = 3,$$

An atomic **formula** $R(t_1, t_2)$ is satisfied if the tuple of values of its arguments, i.e., $\langle \text{Val}_s^{\mathfrak{M}}(t_1), \text{Val}_s^{\mathfrak{M}}(t_2) \rangle$, is **an element** of $R^{\mathfrak{M}}$. So, e.g., we have $\mathfrak{M}, s \models R(b, f(a, b))$ since $\langle \text{Val}_s^{\mathfrak{M}}(b), \text{Val}_s^{\mathfrak{M}}(f(a, b)) \rangle = \langle 2, 3 \rangle \in R^{\mathfrak{M}}$, but $\mathfrak{M} \not\models R(x, f(a, b))$ since $\langle 1, 3 \rangle \notin R^{\mathfrak{M}}[s]$.

To determine if a non-atomic formula φ is satisfied, you apply the clauses in the inductive definition that applies to the main connective. For instance, the main connective in $R(a, a) \rightarrow (R(b, x) \vee R(x, b))$ is the \rightarrow , and

$$\begin{aligned} \mathfrak{M}, s \models R(a, a) \rightarrow (R(b, x) \vee R(x, b)) \text{ iff} \\ \mathfrak{M}, s \not\models R(a, a) \text{ or } \mathfrak{M}, s \models R(b, x) \vee R(x, b) \end{aligned}$$

Since $\mathfrak{M}, s \models R(a, a)$ (because $\langle 1, 1 \rangle \in R^{\mathfrak{M}}$) we can't yet determine the answer and must first figure out if $\mathfrak{M}, s \models R(b, x) \vee R(x, b)$:

$$\begin{aligned} \mathfrak{M}, s \models R(b, x) \vee R(x, b) &\text{ iff} \\ \mathfrak{M}, s \models R(b, x) &\text{ or } \mathfrak{M}, s \models R(x, b) \end{aligned}$$

And this is the case, since $\mathfrak{M}, s \models R(x, b)$ (because $\langle 1, 2 \rangle \in R^{\mathfrak{M}}$).

Recall that an x -variant of s is a variable assignment that differs from s at most in what it assigns to x . For every element of $|\mathfrak{M}|$, there is an x -variant of s : $s_1(x) = 1$, $s_2(x) = 2$, $s_3(x) = 3$, $s_4(x) = 4$, and with $s_i(y) = s(y) = 1$ for all variables y other than x are all the x -variants of s for the structure \mathfrak{M} . Note, in particular, that $s_1 = s$ is also an x -variant of s , i.e., s is an x -variant of itself.

To determine if an existentially quantified formula $\exists x \varphi(x)$ is satisfied, we have to determine if $\mathfrak{M}, s' \models \varphi(x)$ for at least one x -variant s' of s . So,

$$\mathfrak{M}, s \models \exists x (R(b, x) \vee R(x, b)),$$

since $\mathfrak{M}, s_1 \models R(b, x) \vee R(x, b)$ (s_3 would also fit the bill). But,

$$\mathfrak{M}, s \not\models \exists x (R(b, x) \wedge R(x, b))$$

since for none of the s_i , $\mathfrak{M}, s_i \models R(b, x) \wedge R(x, b)$.

To determine if a universally quantified formula $\forall x \varphi(x)$ is satisfied, we have to determine if $\mathfrak{M}, s' \models \varphi(x)$ for all x -variants s' of s . So,

$$\mathfrak{M}, s \models \forall x (R(x, a) \rightarrow R(a, x)),$$

since $\mathfrak{M}, s_i \models R(x, a) \rightarrow R(a, x)$ for all s_i ($\mathfrak{M}, s_1 \models R(a, x)$ and $\mathfrak{M}, s_j \not\models R(a, x)$ for $j = 2, 3$, and 4). But,

$$\mathfrak{M}, s \not\models \forall x (R(a, x) \rightarrow R(x, a))$$

since $\mathfrak{M}, s_2 \not\models R(a, x) \rightarrow R(x, a)$ (because $\mathfrak{M}, s_2 \models R(a, x)$ and $\mathfrak{M}, s_2 \not\models R(x, a)$).

For a more complicated case, consider

$$\forall x (R(a, x) \rightarrow \exists y R(x, y)).$$

Since $\mathfrak{M}, s_3 \not\models R(a, x)$ and $\mathfrak{M}, s_4 \not\models R(a, x)$, the interesting cases where we have to worry about the consequent of the conditional are only s_1 and s_2 . Does $\mathfrak{M}, s_1 \models \exists y R(x, y)$ hold? It does if there is at least one y -variant s'_1 of s_1 so that $\mathfrak{M}, s'_1 \models R(x, y)$. In fact, s_1 is such a y -variant ($s_1(x) = 1$, $s_1(y) = 1$, and $\langle 1, 1 \rangle \in R^{\mathfrak{M}}$), so the answer is yes. To determine if $\mathfrak{M}, s_2 \models \exists y R(x, y)$ we have to look at the y -variants of s_2 . Here, s_2 itself does not satisfy $R(x, y)$ ($s_2(x) = 2$,

$s_2(y) = 1$, and $\langle 2, 1 \rangle \notin R^{\mathfrak{M}}$. However, consider $s'_2 \sim_y s_2$ with $s'_2(y) = 3$. $\mathfrak{M}, s'_2 \models R(x, y)$ since $\langle 2, 3 \rangle \in R^{\mathfrak{M}}$, and so $\mathfrak{M}, s_2 \models \exists y R(x, y)$. In sum, for every x -variant s_i of s , either $\mathfrak{M}, s_i \not\models R(a, x)$ ($i = 3, 4$) or $\mathfrak{M}, s_i \models \exists y R(x, y)$ ($i = 1, 2$), and so

$$\mathfrak{M}, s \models \forall x (R(a, x) \rightarrow \exists y R(x, y)).$$

On the other hand,

$$\mathfrak{M}, s \not\models \exists x (R(a, x) \wedge \forall y R(x, y)).$$

The only x -variants s_i of s with $\mathfrak{M}, s_i \models R(a, x)$ are s_1 and s_2 . But for each, there is in turn a y -variant $s'_i \sim_y s_i$ with $s'_i(y) = 4$ so that $\mathfrak{M}, s'_i \not\models R(x, y)$ and so $\mathfrak{M}, s_i \not\models \forall y R(x, y)$ for $i = 1, 2$. In sum, none of the x -variants $s_i \sim_x s$ are such that $\mathfrak{M}, s_i \models R(a, x) \wedge \forall y R(x, y)$.

Problem 1.5. Let $\mathcal{L} = \{c, f, A\}$ with one **constant symbol**, one **one-place function symbol** and one **two-place predicate symbol**, and let the **structure** \mathfrak{M} be given by

1. $|\mathfrak{M}| = \{1, 2, 3\}$
2. $c^{\mathfrak{M}} = 3$
3. $f^{\mathfrak{M}}(1) = 2, f^{\mathfrak{M}}(2) = 3, f^{\mathfrak{M}}(3) = 2$
4. $A^{\mathfrak{M}} = \{\langle 1, 2 \rangle, \langle 2, 3 \rangle, \langle 3, 3 \rangle\}$

(a) Let $s(v) = 1$ for all variables v . Find out whether

$$\mathfrak{M}, s \models \exists x (A(f(z), c) \rightarrow \forall y (A(y, x) \vee A(f(y), x)))$$

Explain why or why not.

(b) Give a different structure and variable assignment in which the formula is not satisfied.

1.12 Variable Assignments

fol:syn:ass: sec A **variable** assignment s provides a value for *every* variable—and there are explanation infinitely many of them. This is of course not necessary. We require **variable** assignments to assign values to all **variables** simply because it makes things a lot easier. The value of a term t , and whether or not a **formula** φ is satisfied in a **structure** with respect to s , only depend on the assignments s makes to the **variables** in t and the free **variables** of φ . This is the content of the next two propositions. To make the idea of “depends on” precise, we show that any two variable assignments that agree on all the variables in t give the same value, and that φ is satisfied relative to one iff it is satisfied relative to the other if two variable assignments agree on all free variables of φ .

fol:syn:ass: prop:valindep **Proposition 1.36.** *If the **variables** in a term t are among x_1, \dots, x_n , and $s_1(x_i) = s_2(x_i)$ for $i = 1, \dots, n$, then $\text{Val}_{s_1}^{\mathfrak{M}}(t) = \text{Val}_{s_2}^{\mathfrak{M}}(t)$.*

Proof. By induction on the complexity of t . For the base case, t can be a **constant symbol** or one of the variables x_1, \dots, x_n . If $t = c$, then $\text{Val}_{s_1}^{\mathfrak{M}}(t) = c^{\mathfrak{M}} = \text{Val}_{s_2}^{\mathfrak{M}}(t)$. If $t = x_i$, $s_1(x_i) = s_2(x_i)$ by the hypothesis of the proposition, and so $\text{Val}_{s_1}^{\mathfrak{M}}(t) = s_1(x_i) = s_2(x_i) = \text{Val}_{s_2}^{\mathfrak{M}}(t)$.

For the inductive step, assume that $t = f(t_1, \dots, t_k)$ and that the claim holds for t_1, \dots, t_k . Then

$$\begin{aligned} \text{Val}_{s_1}^{\mathfrak{M}}(t) &= \text{Val}_{s_1}^{\mathfrak{M}}(f(t_1, \dots, t_k)) = \\ &= f^{\mathfrak{M}}(\text{Val}_{s_1}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s_1}^{\mathfrak{M}}(t_k)) \end{aligned}$$

For $j = 1, \dots, k$, the **variables** of t_j are among x_1, \dots, x_n . So by induction hypothesis, $\text{Val}_{s_1}^{\mathfrak{M}}(t_j) = \text{Val}_{s_2}^{\mathfrak{M}}(t_j)$. So,

$$\begin{aligned} \text{Val}_{s_1}^{\mathfrak{M}}(t) &= \text{Val}_{s_2}^{\mathfrak{M}}(f(t_1, \dots, t_k)) = \\ &= f^{\mathfrak{M}}(\text{Val}_{s_1}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s_1}^{\mathfrak{M}}(t_k)) = \\ &= f^{\mathfrak{M}}(\text{Val}_{s_2}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s_2}^{\mathfrak{M}}(t_k)) = \\ &= \text{Val}_{s_2}^{\mathfrak{M}}(f(t_1, \dots, t_k)) = \text{Val}_{s_2}^{\mathfrak{M}}(t). \end{aligned}$$

□

Proposition 1.37. *If the free **variables** in φ are among x_1, \dots, x_n , and $s_1(x_i) = s_2(x_i)$ for $i = 1, \dots, n$, then $\mathfrak{M}, s_1 \models \varphi$ iff $\mathfrak{M}, s_2 \models \varphi$.*

*fol:syn:ass:
prop:satind*

Proof. We use induction on the complexity of φ . For the base case, where φ is atomic, φ can be: \top , \perp , $R(t_1, \dots, t_k)$ for a k -place predicate R and terms t_1, \dots, t_k , or $t_1 = t_2$ for terms t_1 and t_2 .

1. $\varphi \equiv \top$: both $\mathfrak{M}, s_1 \models \varphi$ and $\mathfrak{M}, s_2 \models \varphi$.
2. $\varphi \equiv \perp$: both $\mathfrak{M}, s_1 \not\models \varphi$ and $\mathfrak{M}, s_2 \not\models \varphi$.
3. $\varphi \equiv R(t_1, \dots, t_k)$: let $\mathfrak{M}, s_1 \models \varphi$. Then

$$\langle \text{Val}_{s_1}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s_1}^{\mathfrak{M}}(t_k) \rangle \in R^{\mathfrak{M}}.$$

For $i = 1, \dots, k$, $\text{Val}_{s_1}^{\mathfrak{M}}(t_i) = \text{Val}_{s_2}^{\mathfrak{M}}(t_i)$ by [Proposition 1.36](#). So we also have $\langle \text{Val}_{s_2}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s_2}^{\mathfrak{M}}(t_k) \rangle \in R^{\mathfrak{M}}$.

4. $\varphi \equiv t_1 = t_2$: suppose $\mathfrak{M}, s_1 \models \varphi$. Then $\text{Val}_{s_1}^{\mathfrak{M}}(t_1) = \text{Val}_{s_1}^{\mathfrak{M}}(t_2)$. So,

$$\begin{aligned} \text{Val}_{s_2}^{\mathfrak{M}}(t_1) &= \text{Val}_{s_1}^{\mathfrak{M}}(t_1) && \text{(by Proposition 1.36)} \\ &= \text{Val}_{s_1}^{\mathfrak{M}}(t_2) && \text{(since } \mathfrak{M}, s_1 \models t_1 = t_2 \text{)} \\ &= \text{Val}_{s_2}^{\mathfrak{M}}(t_2) && \text{(by Proposition 1.36),} \end{aligned}$$

so $\mathfrak{M}, s_2 \models t_1 = t_2$.

Now assume $\mathfrak{M}, s_1 \models \psi$ iff $\mathfrak{M}, s_2 \models \psi$ for all **formulas** ψ less complex than φ . The induction step proceeds by cases determined by the main operator of φ . In each case, we only demonstrate the forward direction of the **biconditional**; the proof of the reverse direction is symmetrical. In all cases except those for the quantifiers, we apply the induction hypothesis to sub-**formulas** ψ of φ . The free variables of ψ are among those of φ . Thus, if s_1 and s_2 agree on the free variables of φ , they also agree on those of ψ , and the induction hypothesis applies to ψ .

1. $\varphi \equiv \neg\psi$: if $\mathfrak{M}, s_1 \models \varphi$, then $\mathfrak{M}, s_1 \not\models \psi$, so by the induction hypothesis, $\mathfrak{M}, s_2 \not\models \psi$, hence $\mathfrak{M}, s_2 \models \varphi$.
2. $\varphi \equiv \psi \wedge \chi$: if $\mathfrak{M}, s_1 \models \varphi$, then $\mathfrak{M}, s_1 \models \psi$ and $\mathfrak{M}, s_1 \models \chi$, so by induction hypothesis, $\mathfrak{M}, s_2 \models \psi$ and $\mathfrak{M}, s_2 \models \chi$. Hence, $\mathfrak{M}, s_2 \models \varphi$.
3. $\varphi \equiv \psi \vee \chi$: if $\mathfrak{M}, s_1 \models \varphi$, then $\mathfrak{M}, s_1 \models \psi$ or $\mathfrak{M}, s_1 \models \chi$. By induction hypothesis, $\mathfrak{M}, s_2 \models \psi$ or $\mathfrak{M}, s_2 \models \chi$, so $\mathfrak{M}, s_2 \models \varphi$.
4. $\varphi \equiv \psi \rightarrow \chi$: if $\mathfrak{M}, s_1 \models \varphi$, then $\mathfrak{M}, s_1 \not\models \psi$ or $\mathfrak{M}, s_1 \models \chi$. By the induction hypothesis, $\mathfrak{M}, s_2 \not\models \psi$ or $\mathfrak{M}, s_2 \models \chi$, so $\mathfrak{M}, s_2 \models \varphi$.
5. $\varphi \equiv \psi \leftrightarrow \chi$: if $\mathfrak{M}, s_1 \models \varphi$, then either $\mathfrak{M}, s_1 \models \psi$ and $\mathfrak{M}, s_1 \models \chi$, or $\mathfrak{M}, s_1 \not\models \psi$ and $\mathfrak{M}, s_1 \not\models \chi$. By the induction hypothesis, either $\mathfrak{M}, s_2 \models \psi$ and $\mathfrak{M}, s_2 \models \chi$ or $\mathfrak{M}, s_2 \not\models \psi$ and $\mathfrak{M}, s_2 \not\models \chi$. In either case, $\mathfrak{M}, s_2 \models \varphi$.
6. $\varphi \equiv \exists x \psi$: if $\mathfrak{M}, s_1 \models \varphi$, there is an x -variant s'_1 of s_1 so that $\mathfrak{M}, s'_1 \models \psi$. Let s'_2 be the x -variant of s_2 that assigns the same thing to x as does s'_1 . The free variables of ψ are among x_1, \dots, x_n , and x . $s'_1(x_i) = s'_2(x_i)$, since s'_1 and s'_2 are x -variants of s_1 and s_2 , respectively, and by hypothesis $s_1(x_i) = s_2(x_i)$. $s'_1(x) = s'_2(x)$ by the way we have defined s'_2 . Then the induction hypothesis applies to ψ and s'_1, s'_2 , so $\mathfrak{M}, s'_2 \models \psi$. Hence, there is an x -variant of s_2 that satisfies ψ , and so $\mathfrak{M}, s_2 \models \varphi$.
7. $\varphi \equiv \forall x \psi$: if $\mathfrak{M}, s_1 \models \varphi$, then for every x -variant s'_1 of s_1 , $\mathfrak{M}, s'_1 \models \psi$. Take an arbitrary x -variant s'_2 of s_2 , let s'_1 be the x -variant of s_1 which assigns the same thing to x as does s'_2 . The free variables of ψ are among x_1, \dots, x_n , and x . $s'_1(x_i) = s'_2(x_i)$, since s'_1 and s'_2 are x -variants of s_1 and s_2 , respectively, and by hypothesis $s_1(x_i) = s_2(x_i)$. $s'_1(x) = s'_2(x)$ by the way we have defined s'_1 . Then the induction hypothesis applies to ψ and s'_1, s_2 , and we have $\mathfrak{M}, s'_2 \models \psi$. Since s'_2 is an arbitrary x -variant of s_2 , every x -variant of s_2 satisfies ψ , and so $\mathfrak{M}, s_2 \models \varphi$.

By induction, we get that $\mathfrak{M}, s_1 \models \varphi$ iff $\mathfrak{M}, s_2 \models \varphi$ whenever the free **variables** in φ are among x_1, \dots, x_n and $s_1(x_i) = s_2(x_i)$ for $i = 1, \dots, n$. \square

Problem 1.6. Complete the proof of [Proposition 1.37](#).

explanation

Sentences have no free variables, so any two variable assignments assign the same things to all the (zero) free variables of any sentence. The proposition just proved then means that whether or not a **sentence** is satisfied in a structure relative to a variable assignment is completely independent of the assignment. We'll record this fact. It justifies the definition of satisfaction of a **sentence** in a **structure** (without mentioning a variable assignment) that follows.

Corollary 1.38. *If φ is a **sentence** and s a variable assignment, then $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s' \models \varphi$ for every variable assignment s' .* fol:syn:ass: cor:sat-sentence

Proof. Let s' be any variable assignment. Since φ is a **sentence**, it has no free variables, and so every variable assignment s' trivially assigns the same things to all free variables of φ as does s . So the condition of [Proposition 1.37](#) is satisfied, and we have $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}, s' \models \varphi$. \square

Definition 1.39. If φ is a **sentence**, we say that a **structure** \mathfrak{M} *satisfies* φ , $\mathfrak{M} \models \varphi$, iff $\mathfrak{M}, s \models \varphi$ for all variable assignments s . fol:syn:ass: defn:satisfaction

If $\mathfrak{M} \models \varphi$, we also simply say that φ is *true in* \mathfrak{M} .

Proposition 1.40. *Let \mathfrak{M} be a **structure**, φ be a **sentence**, and s a variable assignment. $\mathfrak{M} \models \varphi$ iff $\mathfrak{M}, s \models \varphi$.* fol:syn:ass: prop:sentence-sat-true

Proof. Exercise. \square

Problem 1.7. Prove [Proposition 1.40](#)

Proposition 1.41. *Suppose $\varphi(x)$ only contains x free, and \mathfrak{M} is a **structure**. Then:* fol:syn:ass: prop:sat-quant

1. $\mathfrak{M} \models \exists x \varphi(x)$ iff $\mathfrak{M}, s \models \varphi(x)$ for at least one variable assignment s .
2. $\mathfrak{M} \models \forall x \varphi(x)$ iff $\mathfrak{M}, s \models \varphi(x)$ for all variable assignments s .

Proof. Exercise. \square

Problem 1.8. Prove [Proposition 1.41](#).

Problem 1.9. Suppose \mathcal{L} is a language without **function symbols**. Given a **structure** \mathfrak{M} , c a **constant symbol** and $a \in |\mathfrak{M}|$, define $\mathfrak{M}[a/c]$ to be the **structure** that is just like \mathfrak{M} , except that $c^{\mathfrak{M}[a/c]} = a$. Define $\mathfrak{M} \models \varphi$ for **sentences** φ by:

1. $\varphi \equiv \perp$: not $\mathfrak{M} \models \varphi$.
2. $\varphi \equiv \top$: $\mathfrak{M} \models \varphi$.
3. $\varphi \equiv R(d_1, \dots, d_n)$: $\mathfrak{M} \models \varphi$ iff $\langle d_1^{\mathfrak{M}}, \dots, d_n^{\mathfrak{M}} \rangle \in R^{\mathfrak{M}}$.
4. $\varphi \equiv d_1 = d_2$: $\mathfrak{M} \models \varphi$ iff $d_1^{\mathfrak{M}} = d_2^{\mathfrak{M}}$.
5. $\varphi \equiv \neg\psi$: $\mathfrak{M} \models \varphi$ iff not $\mathfrak{M} \models \psi$.

6. $\varphi \equiv (\psi \wedge \chi)$: $\mathfrak{M} \models \varphi$ iff $\mathfrak{M} \models \psi$ and $\mathfrak{M} \models \chi$.
7. $\varphi \equiv (\psi \vee \chi)$: $\mathfrak{M} \models \varphi$ iff $\mathfrak{M} \models \psi$ or $\mathfrak{M} \models \chi$ (or both).
8. $\varphi \equiv (\psi \rightarrow \chi)$: $\mathfrak{M} \models \varphi$ iff not $\mathfrak{M} \models \psi$ or $\mathfrak{M} \models \chi$ (or both).
9. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\mathfrak{M} \models \varphi$ iff either both $\mathfrak{M} \models \psi$ and $\mathfrak{M} \models \chi$, or neither $\mathfrak{M} \models \psi$ nor $\mathfrak{M} \models \chi$.
10. $\varphi \equiv \forall x \psi$: $\mathfrak{M} \models \varphi$ iff for all $a \in |\mathfrak{M}|$, $\mathfrak{M}[a/c] \models \psi[c/x]$, if c does not occur in ψ .
11. $\varphi \equiv \exists x \psi$: $\mathfrak{M} \models \varphi$ iff there is an $a \in |\mathfrak{M}|$ such that $\mathfrak{M}[a/c] \models \psi[c/x]$, if c does not occur in ψ .

Let x_1, \dots, x_n be all free variables in φ , c_1, \dots, c_n constant symbols not in φ , $a_1, \dots, a_n \in |\mathfrak{M}|$, and $s(x_i) = a_i$.

Show that $\mathfrak{M}, s \models \varphi$ iff $\mathfrak{M}[a_1/c_1, \dots, a_n/c_n] \models \varphi[c_1/x_1] \dots [c_n/x_n]$.

(This problem shows that it is possible to give a semantics for first-order logic that makes do without variable assignments.)

Problem 1.10. Suppose that f is a function symbol not in $\varphi(x, y)$. Show that there is a structure \mathfrak{M} such that $\mathfrak{M} \models \forall x \exists y \varphi(x, y)$ iff there is an \mathfrak{M}' such that $\mathfrak{M}' \models \forall x \varphi(x, f(x))$.

(This problem is a special case of what's known as Skolem's Theorem; $\forall x \varphi(x, f(x))$ is called a *Skolem normal form* of $\forall x \exists y \varphi(x, y)$.)

1.13 Extensionality

fol:syn:ext:
sec

Extensionality, sometimes called relevance, can be expressed informally as follows: the only thing that bears upon the satisfaction of formula φ in a structure \mathfrak{M} relative to a variable assignment s , are the assignments made by \mathfrak{M} and s to the elements of the language that actually appear in φ . explanation

One immediate consequence of extensionality is that where two structures \mathfrak{M} and \mathfrak{M}' agree on all the elements of the language appearing in a sentence φ and have the same domain, \mathfrak{M} and \mathfrak{M}' must also agree on whether or not φ itself is true.

fol:syn:ext:
prop:extensionality

Proposition 1.42 (Extensionality). *Let φ be a formula, and \mathfrak{M}_1 and \mathfrak{M}_2 be structures with $|\mathfrak{M}_1| = |\mathfrak{M}_2|$, and s a variable assignment on $|\mathfrak{M}_1| = |\mathfrak{M}_2|$. If $c^{\mathfrak{M}_1} = c^{\mathfrak{M}_2}$, $R^{\mathfrak{M}_1} = R^{\mathfrak{M}_2}$, and $f^{\mathfrak{M}_1} = f^{\mathfrak{M}_2}$ for every constant symbol c , relation symbol R , and function symbol f occurring in φ , then $\mathfrak{M}_1, s \models \varphi$ iff $\mathfrak{M}_2, s \models \varphi$.*

Proof. First prove (by induction on t) that for every term, $\text{Val}_s^{\mathfrak{M}_1}(t) = \text{Val}_s^{\mathfrak{M}_2}(t)$. Then prove the proposition by induction on φ , making use of the claim just proved for the induction basis (where φ is atomic). \square

Problem 1.11. Carry out the proof of Proposition 1.42 in detail.

Corollary 1.43 (Extensionality for Sentences). *Let φ be a sentence and $\mathfrak{M}_1, \mathfrak{M}_2$ as in Proposition 1.42. Then $\mathfrak{M}_1 \models \varphi$ iff $\mathfrak{M}_2 \models \varphi$.* fol:syn:ext:
cor:extensionality-sent

Proof. Follows from Proposition 1.42 by Corollary 1.38. □

Moreover, the value of a term, and whether or not a structure satisfies a formula, only depends on the values of its subterms.

Proposition 1.44. *Let \mathfrak{M} be a structure, t and t' terms, and s a variable assignment. Let $s' \sim_x s$ be the x -variant of s given by $s'(x) = \text{Val}_s^{\mathfrak{M}}(t')$. Then $\text{Val}_s^{\mathfrak{M}}(t[t'/x]) = \text{Val}_{s'}^{\mathfrak{M}}(t)$.* fol:syn:ext:
prop:ext-terms

Proof. By induction on t .

1. If t is a constant, say, $t \equiv c$, then $t[t'/x] = c$, and $\text{Val}_s^{\mathfrak{M}}(c) = c^{\mathfrak{M}} = \text{Val}_{s'}^{\mathfrak{M}}(c)$.
2. If t is a variable other than x , say, $t \equiv y$, then $t[t'/x] = y$, and $\text{Val}_s^{\mathfrak{M}}(y) = \text{Val}_{s'}^{\mathfrak{M}}(y)$ since $s' \sim_x s$.
3. If $t \equiv x$, then $t[t'/x] = t'$. But $\text{Val}_s^{\mathfrak{M}}(x) = \text{Val}_s^{\mathfrak{M}}(t')$ by definition of s' .
4. If $t \equiv f(t_1, \dots, t_n)$ then we have:

$$\begin{aligned}
 \text{Val}_s^{\mathfrak{M}}(t[t'/x]) &= \\
 &= \text{Val}_s^{\mathfrak{M}}(f(t_1[t'/x], \dots, t_n[t'/x])) \\
 &\quad \text{by definition of } t[t'/x] \\
 &= f^{\mathfrak{M}}(\text{Val}_s^{\mathfrak{M}}(t_1[t'/x]), \dots, \text{Val}_s^{\mathfrak{M}}(t_n[t'/x])) \\
 &\quad \text{by definition of } \text{Val}_s^{\mathfrak{M}}(f(\dots)) \\
 &= f^{\mathfrak{M}}(\text{Val}_{s'}^{\mathfrak{M}}(t_1), \dots, \text{Val}_{s'}^{\mathfrak{M}}(t_n)) \\
 &\quad \text{by induction hypothesis} \\
 &= \text{Val}_{s'}^{\mathfrak{M}}(t) \text{ by definition of } \text{Val}_{s'}^{\mathfrak{M}}(f(\dots))
 \end{aligned}$$

□

Proposition 1.45. *Let \mathfrak{M} be a structure, φ a formula, t a term, and s a variable assignment. Let $s' \sim_x s$ be the x -variant of s given by $s'(x) = \text{Val}_s^{\mathfrak{M}}(t)$. Then $\mathfrak{M}, s \models \varphi[t/x]$ iff $\mathfrak{M}, s' \models \varphi$.* fol:syn:ext:
prop:ext-formulas

Proof. Exercise. □

Problem 1.12. Prove Proposition 1.45

1.14 Semantic Notions

fol:syn:sem:
sec

Give the definition of **structures** for first-order languages, we can define some basic semantic properties of and relationships between sentences. The simplest of these is the notion of *validity* of a sentence. A sentence is valid if it is satisfied in every **structure**. Valid sentences are those that are satisfied regardless of how the non-logical symbols in it are interpreted. Valid sentences are therefore also called *logical truths*—they are true, i.e., satisfied, in any **structure** and hence their truth depends only on the logical symbols occurring in them and their syntactic **structure**, but not on the non-logical symbols or their interpretation.

explanation

Definition 1.46 (Validity). A sentence φ is *valid*, $\vDash \varphi$, iff $\mathfrak{M} \models \varphi$ for every **structure** \mathfrak{M} .

Definition 1.47 (Entailment). A set of sentences Γ *entails* a sentence φ , $\Gamma \vDash \varphi$, iff for every **structure** \mathfrak{M} with $\mathfrak{M} \models \Gamma$, $\mathfrak{M} \models \varphi$.

Definition 1.48 (Satisfiability). A set of sentences Γ is *satisfiable* if $\mathfrak{M} \models \Gamma$ for some **structure** \mathfrak{M} . If Γ is not satisfiable it is called *unsatisfiable*.

Proposition 1.49. *A sentence φ is valid iff $\Gamma \vDash \varphi$ for every set of sentences Γ .*

Proof. For the forward direction, let φ be valid, and let Γ be a set of sentences. Let \mathfrak{M} be a **structure** so that $\mathfrak{M} \models \Gamma$. Since φ is valid, $\mathfrak{M} \models \varphi$, hence $\Gamma \vDash \varphi$.

For the contrapositive of the reverse direction, let φ be invalid, so there is a **structure** \mathfrak{M} with $\mathfrak{M} \not\models \varphi$. When $\Gamma = \{\top\}$, since \top is valid, $\mathfrak{M} \models \Gamma$. Hence, there is a **structure** \mathfrak{M} so that $\mathfrak{M} \models \Gamma$ but $\mathfrak{M} \not\models \varphi$, hence Γ does not entail φ . \square

fol:syn:sem:
prop:entails-unsat

Proposition 1.50. *$\Gamma \vDash \varphi$ iff $\Gamma \cup \{\neg\varphi\}$ is unsatisfiable.*

Proof. For the forward direction, suppose $\Gamma \vDash \varphi$ and suppose to the contrary that there is a **structure** \mathfrak{M} so that $\mathfrak{M} \models \Gamma \cup \{\neg\varphi\}$. Since $\mathfrak{M} \models \Gamma$ and $\Gamma \vDash \varphi$, $\mathfrak{M} \models \varphi$. Also, since $\mathfrak{M} \models \Gamma \cup \{\neg\varphi\}$, $\mathfrak{M} \models \neg\varphi$, so we have both $\mathfrak{M} \models \varphi$ and $\mathfrak{M} \models \neg\varphi$, a contradiction. Hence, there can be no such **structure** \mathfrak{M} , so $\Gamma \cup \{\varphi\}$ is unsatisfiable.

For the reverse direction, suppose $\Gamma \cup \{\neg\varphi\}$ is unsatisfiable. So for every **structure** \mathfrak{M} , either $\mathfrak{M} \not\models \Gamma$ or $\mathfrak{M} \models \varphi$. Hence, for every **structure** \mathfrak{M} with $\mathfrak{M} \models \Gamma$, $\mathfrak{M} \models \varphi$, so $\Gamma \vDash \varphi$. \square

Problem 1.13. 1. Show that $\Gamma \vDash \perp$ iff Γ is unsatisfiable.

2. Show that $\Gamma \cup \{\varphi\} \vDash \perp$ iff $\Gamma \vDash \neg\varphi$.

3. Suppose c does not occur in φ or Γ . Show that $\Gamma \vDash \forall x \varphi$ iff $\Gamma \vDash \varphi[c/x]$.

Proposition 1.51. *If $\Gamma \subseteq \Gamma'$ and $\Gamma \vDash \varphi$, then $\Gamma' \vDash \varphi$.*

Proof. Suppose that $\Gamma \subseteq \Gamma'$ and $\Gamma \models \varphi$. Let \mathfrak{M} be such that $\mathfrak{M} \models \Gamma'$; then $\mathfrak{M} \models \Gamma$, and since $\Gamma \models \varphi$, we get that $\mathfrak{M} \models \varphi$. Hence, whenever $\mathfrak{M} \models \Gamma'$, $\mathfrak{M} \models \varphi$, so $\Gamma' \models \varphi$. \square

Theorem 1.52 (Semantic Deduction Theorem). $\Gamma \cup \{\varphi\} \models \psi$ iff $\Gamma \models \varphi \rightarrow \psi$. *fol:syn:sem:
thm:sem-deduction*

Proof. For the forward direction, let $\Gamma \cup \{\varphi\} \models \psi$ and let \mathfrak{M} be a **structure** so that $\mathfrak{M} \models \Gamma$. If $\mathfrak{M} \models \varphi$, then $\mathfrak{M} \models \Gamma \cup \{\varphi\}$, so since $\Gamma \cup \{\varphi\}$ entails ψ , we get $\mathfrak{M} \models \psi$. Therefore, $\mathfrak{M} \models \varphi \rightarrow \psi$, so $\Gamma \models \varphi \rightarrow \psi$.

For the reverse direction, let $\Gamma \models \varphi \rightarrow \psi$ and \mathfrak{M} be a **structure** so that $\mathfrak{M} \models \Gamma \cup \{\varphi\}$. Then $\mathfrak{M} \models \Gamma$, so $\mathfrak{M} \models \varphi \rightarrow \psi$, and since $\mathfrak{M} \models \varphi$, $\mathfrak{M} \models \psi$. Hence, whenever $\mathfrak{M} \models \Gamma \cup \{\varphi\}$, $\mathfrak{M} \models \psi$, so $\Gamma \cup \{\varphi\} \models \psi$. \square

Proposition 1.53. Let \mathfrak{M} be a **structure**, and $\varphi(x)$ a **formula** with one free variable x , and t a closed term. Then: *fol:syn:sem:
prop:quant-terms*

1. $\varphi(t) \models \exists x \varphi(x)$
2. $\forall x \varphi(x) \models \varphi(t)$

Proof. 1. Suppose $\mathfrak{M} \models \varphi(t)$. Let s be a variable assignment with $s(x) = \text{Val}^{\mathfrak{M}}(t)$. Then $\mathfrak{M}, s \models \varphi(t)$ since $\varphi(t)$ is a **sentence**. By [Proposition 1.45](#), $\mathfrak{M}, s \models \varphi(x)$. By [Proposition 1.41](#), $\mathfrak{M} \models \exists x \varphi(x)$.

2. Suppose $\mathfrak{M} \models \forall x \varphi(x)$. Let s be a variable assignment with $s(x) = \text{Val}^{\mathfrak{M}}(t)$. By [Proposition 1.41](#), $\mathfrak{M}, s \models \varphi(x)$. By [Proposition 1.45](#), $\mathfrak{M}, s \models \varphi(t)$. By [Proposition 1.40](#), $\mathfrak{M} \models \varphi(t)$ since $\varphi(t)$ is a **sentence**. \square

Problem 1.14. Complete the proof of [Proposition 1.53](#).

Chapter 2

Theories and Their Models

2.1 Introduction

fol:mat:int:
sec

The development of the axiomatic method is a significant achievement in the history of science, and is of special importance in the history of mathematics. An axiomatic development of a field involves the clarification of many questions: What is the field about? What are the most fundamental concepts? How are they related? Can all the concepts of the field be defined in terms of these fundamental concepts? What laws do, and must, these concepts obey?

explanation

The axiomatic method and logic were made for each other. Formal logic provides the tools for formulating axiomatic theories, for proving theorems from the axioms of the theory in a precisely specified way, for studying the properties of all systems satisfying the axioms in a systematic way.

Definition 2.1. A set of sentences Γ is *closed* iff, whenever $\Gamma \models \varphi$ then $\varphi \in \Gamma$. The *closure* of a set of sentences Γ is $\{\varphi : \Gamma \models \varphi\}$.

We say that Γ is *axiomatized by* a set of sentences Δ if Γ is the closure of Δ .

We can think of an axiomatic theory as the set of sentences that is axiomatized by its set of axioms Δ . In other words, when we have a first-order language which contains non-logical symbols for the primitives of the axiomatically developed science we wish to study, together with a set of sentences that express the fundamental laws of the science, we can think of the theory as represented by all the sentences in this language that are entailed by the axioms. This ranges from simple examples with only a single primitive and simple axioms, such as the theory of partial orders, to complex theories such as Newtonian mechanics.

explanation

The important logical facts that make this formal approach to the axiomatic method so important are the following. Suppose Γ is an axiom system for a theory, i.e., a set of sentences.

1. We can state precisely when an axiom system captures an intended class of structures. That is, if we are interested in a certain class of structures

tures, we will successfully capture that class by an axiom system Γ iff the structures are exactly those \mathfrak{M} such that $\mathfrak{M} \models \Gamma$.

2. We may fail in this respect because there are \mathfrak{M} such that $\mathfrak{M} \models \Gamma$, but \mathfrak{M} is not one of the structures we intend. This may lead us to add axioms which are not true in \mathfrak{M} .
3. If we are successful at least in the respect that Γ is true in all the intended structures, then a sentence φ is true in all intended structures whenever $\Gamma \models \varphi$. Thus we can use logical tools (such as proof methods) to show that sentences are true in all intended structures simply by showing that they are entailed by the axioms.
4. Sometimes we don't have intended structures in mind, but instead start from the axioms themselves: we begin with some primitives that we want to satisfy certain laws which we codify in an axiom system. One thing that we would like to verify right away is that the axioms do not contradict each other: if they do, there can be no concepts that obey these laws, and we have tried to set up an incoherent theory. We can verify that this doesn't happen by finding a model of Γ . And if there are models of our theory, we can use logical methods to investigate them, and we can also use logical methods to construct models.
5. The independence of the axioms is likewise an important question. It may happen that one of the axioms is actually a consequence of the others, and so is redundant. We can prove that an axiom φ in Γ is redundant by proving $\Gamma \setminus \{\varphi\} \models \varphi$. We can also prove that an axiom is not redundant by showing that $(\Gamma \setminus \{\varphi\}) \cup \{\neg\varphi\}$ is satisfiable. For instance, this is how it was shown that the parallel postulate is independent of the other axioms of geometry.
6. Another important question is that of definability of concepts in a theory: The choice of the language determines what the models of a theory consists of. But not every aspect of a theory must be represented separately in its models. For instance, every ordering \leq determines a corresponding strict ordering $<$ —given one, we can define the other. So it is not necessary that a model of a theory involving such an order must *also* contain the corresponding strict ordering. When is it the case, in general, that one relation can be defined in terms of others? When is it impossible to define a relation in terms of other (and hence must add it to the primitives of the language)?

2.2 Expressing Properties of Structures

explanation It is often useful and important to express conditions on functions and relations, or more generally, that the functions and relations in a structure satisfy these conditions. For instance, we would like to have ways of distinguishing fol:mat:exs:sec

those **structures** for a language which “capture” what we want the **predicate symbols** to “mean” from those that do not. Of course we’re completely free to specify which **structures** we “intend,” e.g., we can specify that the interpretation of the **predicate symbol** \leq must be an ordering, or that we are only interested in interpretations of \mathcal{L} in which the domain consists of sets and \in is interpreted by the “is an element of” relation. But can we do this with **sentences** of the language? In other words, which conditions on a **structure** \mathfrak{M} can we express by a **sentence** (or perhaps a set of **sentences**) in the language of \mathfrak{M} ? There are some conditions that we will not be able to express. For instance, there is no sentence of \mathcal{L}_A which is only true in a **structure** \mathfrak{M} if $|\mathfrak{M}| = \mathbb{N}$. We cannot express “the domain contains only natural numbers.” But there are “structural properties” of **structures** that we perhaps can express. Which properties of **structures** can we express by **sentences**? Or, to put it another way, which collections of **structures** can we describe as those making a **sentence** (or set of **sentences**) true?

Definition 2.2 (Model of a set). Let Γ be a set of **sentences** in a language \mathcal{L} . We say that a **structure** \mathfrak{M} is a model of Γ if $\mathfrak{M} \models \varphi$ for all $\varphi \in \Gamma$.

Example 2.3. The sentence $\forall x x \leq x$ is true in \mathfrak{M} iff $\leq^{\mathfrak{M}}$ is a reflexive relation. The sentence $\forall x \forall y ((x \leq y \wedge y \leq x) \rightarrow x = y)$ is true in \mathfrak{M} iff $\leq^{\mathfrak{M}}$ is anti-symmetric. The sentence $\forall x \forall y \forall z ((x \leq y \wedge y \leq z) \rightarrow x \leq z)$ is true in \mathfrak{M} iff $\leq^{\mathfrak{M}}$ is transitive. Thus, the models of

$$\left\{ \begin{array}{l} \forall x x \leq x, \\ \forall x \forall y ((x \leq y \wedge y \leq x) \rightarrow x = y), \\ \forall x \forall y \forall z ((x \leq y \wedge y \leq z) \rightarrow x \leq z) \end{array} \right\}$$

are exactly those structures in which $\leq^{\mathfrak{M}}$ is reflexive, anti-symmetric, and transitive, i.e., a partial order. Hence, we can take them as axioms for the *first-order theory of partial orders*.

2.3 Examples of First-Order Theories

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Example 2.4. The theory of strict linear orders in the language $\mathcal{L}_<$ is axiomatized by the set

$$\left\{ \begin{array}{l} \forall x \neg x < x, \\ \forall x \forall y ((x < y \vee y < x) \vee x = y), \\ \forall x \forall y \forall z ((x < y \wedge y < z) \rightarrow x < z) \end{array} \right\}$$

It completely captures the intended **structures**: every strict linear order is a model of this axiom system, and vice versa, if R is a linear order on a set X , then the structure \mathfrak{M} with $|\mathfrak{M}| = X$ and $<^{\mathfrak{M}} = R$ is a model of this theory.

Example 2.5. The theory of groups in the language $\mathbf{1}$ (constant symbol), \cdot (two-place function symbol) is axiomatized by

$$\begin{aligned}\forall x (x \cdot \mathbf{1}) &= x \\ \forall x \forall y \forall z (x \cdot (y \cdot z)) &= ((x \cdot y) \cdot z) \\ \forall x \exists y (x \cdot y) &= \mathbf{1}\end{aligned}$$

Example 2.6. The theory of Peano arithmetic is axiomatized by the following sentences in the language of arithmetic \mathcal{L}_A .

$$\begin{aligned}\neg \exists x x' &= \mathbf{0} \\ \forall x \forall y (x' = y' \rightarrow x = y) \\ \forall x \forall y (x < y \leftrightarrow \exists z (x + z' = y)) \\ \forall x (x + \mathbf{0}) &= x \\ \forall x \forall y (x + y') &= (x + y)' \\ \forall x (x \times \mathbf{0}) &= \mathbf{0} \\ \forall x \forall y (x \times y') &= ((x \times y) + x)\end{aligned}$$

plus all sentences of the form

$$(\varphi(\mathbf{0}) \wedge \forall x (\varphi(x) \rightarrow \varphi(x'))) \rightarrow \forall x \varphi(x)$$

Since there are infinitely many sentences of the latter form, this axiom system is infinite. The latter form is called the *induction schema*. (Actually, the induction schema is a bit more complicated than we let on here.)

The third axiom is an *explicit definition* of $<$.

Example 2.7. The theory of pure sets plays an important role in the foundations (and in the philosophy) of mathematics. A set is pure if all its elements are also pure sets. The empty set counts therefore as pure, but a set that has something as an element that is not a set would not be pure. So the pure sets are those that are formed just from the empty set and no “urelements,” i.e., objects that are not themselves sets.

The following might be considered as an axiom system for a theory of pure sets:

$$\begin{aligned}\exists x \neg \exists y y \in x \\ \forall x \forall y (\forall z (z \in x \leftrightarrow z \in y) \rightarrow x = y) \\ \forall x \forall y \exists z \forall u (u \in z \leftrightarrow (u = x \vee u = y)) \\ \forall x \exists y \forall z (z \in y \leftrightarrow \exists u (z \in u \wedge u \in x))\end{aligned}$$

plus all sentences of the form

$$\exists x \forall y (y \in x \leftrightarrow \varphi(y))$$

The first axiom says that there is a set with no **elements** (i.e., \emptyset exists); the second says that sets are extensional; the third that for any sets X and Y , the set $\{X, Y\}$ exists; the fourth that for any sets X and Y , the set $X \cup Y$ exists.

The **sentences** mentioned last are collectively called the *naive comprehension scheme*. It essentially says that for every $\varphi(x)$, the set $\{x : \varphi(x)\}$ exists—so at first glance a true, useful, and perhaps even necessary axiom. It is called “naive” because, as it turns out, it makes this theory unsatisfiable: if you take $\varphi(y)$ to be $\neg y \in y$, you get the **sentence**

$$\exists x \forall y (y \in x \leftrightarrow \neg y \in y)$$

and this **sentence** is not satisfied in any **structure**.

Example 2.8. In the area of *mereology*, the relation of *parthood* is a fundamental relation. Just like theories of sets, there are theories of parthood that axiomatize various conceptions (sometimes conflicting) of this relation.

The language of mereology contains a single two-place predicate symbol P , and $P(x, y)$ “means” that x is a part of y . When we have this interpretation in mind, a **structure** for this language is called a *parthood structure*. Of course, not every structure for a single two-place predicate will really deserve this name. To have a chance of capturing “parthood,” P^M must satisfy some conditions, which we can lay down as axioms for a theory of parthood. For instance, parthood is a partial order on objects: every object is a part (albeit an *improper* part) of itself; no two different objects can be parts of each other; a part of a part of an object is itself part of that object. Note that in this sense “is a part of” resembles “is a subset of,” but does not resemble “is an element of” which is neither reflexive nor transitive.

$$\begin{aligned} &\forall x P(x, x), \\ &\forall x \forall y ((P(x, y) \wedge P(y, x)) \rightarrow x = y), \\ &\forall x \forall y \forall z ((P(x, y) \wedge P(y, z)) \rightarrow P(x, z)), \end{aligned}$$

Moreover, any two objects have a mereological sum (an object that has these two objects as parts, and is minimal in this respect).

$$\forall x \forall y \exists z \forall u (P(z, u) \leftrightarrow (P(x, u) \wedge P(y, u)))$$

These are only some of the basic principles of parthood considered by metaphysicians. Further principles, however, quickly become hard to formulate or write down without first introducing some defined relations. For instance, most metaphysicians interested in mereology also view the following as a valid principle: whenever an object x has a proper part y , it also has a part z that has no parts in common with y , and so that the fusion of y and z is x .

2.4 Expressing Relations in a Structure

One main use **formulas** can be put to is to express properties and relations in a **structure** \mathfrak{M} in terms of the primitives of the language \mathcal{L} of \mathfrak{M} . By this we mean the following: the **domain** of \mathfrak{M} is a set of objects. The **constant symbols**, **function symbols**, and **predicate symbols** are interpreted in \mathfrak{M} by some objects in $|\mathfrak{M}|$, functions on $|\mathfrak{M}|$, and relations on $|\mathfrak{M}|$. For instance, if A_0^2 is in \mathcal{L} , then \mathfrak{M} assigns to it a relation $R = A_0^{2\mathfrak{M}}$. Then the formula $A_0^2(v_1, v_2)$ *expresses* that very relation, in the following sense: if a variable assignment s maps v_1 to $a \in |\mathfrak{M}|$ and v_2 to $b \in |\mathfrak{M}|$, then

$$Rab \quad \text{iff} \quad \mathfrak{M}, s \models A_0^2(v_1, v_2).$$

Note that we have to involve variable assignments here: we can't just say " Rab iff $\mathfrak{M} \models A_0^2(a, b)$ " because a and b are not symbols of our language: they are **elements** of $|\mathfrak{M}|$.

Since we don't just have atomic **formulas**, but can combine them using the logical connectives and the quantifiers, more complex **formulas** can define other relations which aren't directly built into \mathfrak{M} . We're interested in how to do that, and specifically, which relations we can define in a **structure**.

Definition 2.9. Let $\varphi(v_1, \dots, v_n)$ be a **formula** of \mathcal{L} in which only v_1, \dots, v_n occur free, and let \mathfrak{M} be a **structure** for \mathcal{L} . $\varphi(v_1, \dots, v_n)$ *expresses the relation* $R \subseteq |\mathfrak{M}|^n$ iff

$$Ra_1 \dots a_n \quad \text{iff} \quad \mathfrak{M}, s \models \varphi(v_1, \dots, v_n)$$

for any variable assignment s with $s(v_i) = a_i$ ($i = 1, \dots, n$).

Example 2.10. In the standard model of arithmetic \mathfrak{N} , the **formula** $v_1 < v_2 \vee v_1 = v_2$ expresses the \leq relation on \mathbb{N} . The **formula** $v_2 = v_1'$ expresses the successor relation, i.e., the relation $R \subseteq \mathbb{N}^2$ where Rnm holds if m is the successor of n . The formula $v_1 = v_2'$ expresses the predecessor relation. The **formulas** $\exists v_3 (v_3 \neq 0 \wedge v_2 = (v_1 + v_3))$ and $\exists v_3 (v_1 + v_3' = v_2)$ both express the $<$ relation. This means that the predicate symbol $<$ is actually superfluous in the language of arithmetic; it can be defined.

explanation

This idea is not just interesting in specific **structures**, but generally whenever we use a language to describe an intended model or models, i.e., when we consider theories. These theories often only contain a few **predicate symbols** as basic symbols, but in the domain they are used to describe often many other relations play an important role. If these other relations can be systematically expressed by the relations that interpret the basic **predicate symbols** of the language, we say we can *define* them in the language.

Problem 2.1. Find **formulas** in \mathcal{L}_A which define the following relations:

1. n is between i and j ;
2. n evenly divides m (i.e., m is a multiple of n);

3. n is a prime number (i.e., no number other than 1 and n evenly divides n).

Problem 2.2. Suppose the formula $\varphi(v_1, v_2)$ expresses the relation $R \subseteq |\mathfrak{M}|^2$ in a structure \mathfrak{M} . Find formulas that express the following relations:

1. the inverse R^{-1} of R ;
2. the relative product $R \mid R$;

Can you find a way to express R^+ , the transitive closure of R ?

Problem 2.3. Let \mathcal{L} be the language containing a 2-place predicate symbol $<$ only (no other constant symbols, function symbols or predicate symbols—except of course $=$). Let \mathfrak{N} be the structure such that $|\mathfrak{N}| = \mathbb{N}$, and $<^{\mathfrak{N}} = \{\langle n, m \rangle : n < m\}$. Prove the following:

1. $\{0\}$ is definable in \mathfrak{N} ;
2. $\{1\}$ is definable in \mathfrak{N} ;
3. $\{2\}$ is definable in \mathfrak{N} ;
4. for each $n \in \mathbb{N}$, the set $\{n\}$ is definable in \mathfrak{N} ;
5. every finite subset of $|\mathfrak{N}|$ is definable in \mathfrak{N} ;
6. every co-finite subset of $|\mathfrak{N}|$ is definable in \mathfrak{N} (where $X \subseteq \mathbb{N}$ is co-finite iff $\mathbb{N} \setminus X$ is finite).

2.5 The Theory of Sets

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Almost all of mathematics can be developed in the theory of sets. Developing mathematics in this theory involves a number of things. First, it requires a set of axioms for the relation \in . A number of different axiom systems have been developed, sometimes with conflicting properties of \in . The axiom system known as **ZFC**, Zermelo-Fraenkel set theory with the axiom of choice stands out: it is by far the most widely used and studied, because it turns out that its axioms suffice to prove almost all the things mathematicians expect to be able to prove. But before that can be established, it first is necessary to make clear how we can even *express* all the things mathematicians would like to express. For starters, the language contains no constant symbols or function symbols, so it seems at first glance unclear that we can talk about particular sets (such as \emptyset or \mathbb{N}), can talk about operations on sets (such as $X \cup Y$ and $\wp(X)$), let alone other constructions which involve things other than sets, such as relations and functions.

To begin with, “is an element of” is not the only relation we are interested in: “is a subset of” seems almost as important. But we can *define* “is a subset of” in terms of “is an element of.” To do this, we have to find a formula $\varphi(x, y)$ in the language of set theory which is satisfied by a pair of sets $\langle X, Y \rangle$ iff

$X \subseteq Y$. But X is a subset of Y just in case all **elements** of X are also **elements** of Y . So we can define \subseteq by the formula

$$\forall z (z \in x \rightarrow z \in y)$$

Now, whenever we want to use the relation \subseteq in a formula, we could instead use that formula (with x and y suitably replaced, and the bound variable z renamed if necessary). For instance, extensionality of sets means that if any sets x and y are contained in each other, then x and y must be the same set. This can be expressed by $\forall x \forall y ((x \subseteq y \wedge y \subseteq x) \rightarrow x = y)$, or, if we replace \subseteq by the above definition, by

$$\forall x \forall y ((\forall z (z \in x \rightarrow z \in y) \wedge \forall z (z \in y \rightarrow z \in x)) \rightarrow x = y).$$

This is in fact one of the axioms of **ZFC**, the “axiom of extensionality.”

There is no **constant symbol** for \emptyset , but we can express “ x is empty” by $\neg \exists y y \in x$. Then “ \emptyset exists” becomes the **sentence** $\exists x \neg \exists y y \in x$. This is another axiom of **ZFC**. (Note that the axiom of extensionality implies that there is only one empty set.) Whenever we want to talk about \emptyset in the language of set theory, we would write this as “there is a set that’s empty and ...” As an example, to express the fact that \emptyset is a subset of every set, we could write

$$\exists x (\neg \exists y y \in x \wedge \forall z x \subseteq z)$$

where, of course, $x \subseteq z$ would in turn have to be replaced by its definition.

To talk about operations on sets, such as $X \cup Y$ and $\wp(X)$, we have to use a similar trick. There are no function symbols in the language of set theory, but we can express the functional relations $X \cup Y = Z$ and $\wp(X) = Y$ by

$$\begin{aligned} \forall u ((u \in x \vee u \in y) \leftrightarrow u \in z) \\ \forall u (u \subseteq x \leftrightarrow u \in y) \end{aligned}$$

since the **elements** of $X \cup Y$ are exactly the sets that are either **elements** of X or **elements** of Y , and the **elements** of $\wp(X)$ are exactly the subsets of X . However, this doesn’t allow us to use $x \cup y$ or $\wp(x)$ as if they were terms: we can only use the entire **formulas** that define the relations $X \cup Y = Z$ and $\wp(X) = Y$. In fact, we do not know that these relations are ever satisfied, i.e., we do not know that unions and power sets always exist. For instance, the **sentence** $\forall x \exists y \wp(x) = y$ is another axiom of **ZFC** (the power set axiom).

Now what about talk of ordered pairs or functions? Here we have to explain how we can think of ordered pairs and functions as special kinds of sets. One way to define the ordered pair $\langle x, y \rangle$ is as the set $\{\{x\}, \{x, y\}\}$. But like before, we cannot introduce a **function symbol** that names this set; we can only define the relation $\langle x, y \rangle = z$, i.e., $\{\{x\}, \{x, y\}\} = z$:

$$\forall u (u \in z \leftrightarrow (\forall v (v \in u \leftrightarrow v = x) \vee \forall v (v \in u \leftrightarrow (v = x \vee v = y))))$$

This says that the **elements** u of z are exactly those sets which either have x as its only **element** or have x and y as its only **elements** (in other words, those

sets that are either identical to $\{x\}$ or identical to $\{x, y\}$). Once we have this, we can say further things, e.g., that $X \times Y = Z$:

$$\forall z (z \in Z \leftrightarrow \exists x \exists y (x \in X \wedge y \in Y \wedge \langle x, y \rangle = z))$$

A function $f: X \rightarrow Y$ can be thought of as the relation $f(x) = y$, i.e., as the set of pairs $\{\langle x, y \rangle : f(x) = y\}$. We can then say that a set f is a function from X to Y if (a) it is a relation $\subseteq X \times Y$, (b) it is total, i.e., for all $x \in X$ there is some $y \in Y$ such that $\langle x, y \rangle \in f$ and (c) it is functional, i.e., whenever $\langle x, y \rangle, \langle x, y' \rangle \in f$, $y = y'$ (because values of functions must be unique). So “ f is a function from X to Y ” can be written as:

$$\begin{aligned} & \forall u (u \in f \rightarrow \exists x \exists y (x \in X \wedge y \in Y \wedge \langle x, y \rangle = u)) \wedge \\ & \forall x (x \in X \rightarrow (\exists y (y \in Y \wedge \text{maps}(f, x, y)) \wedge \\ & \quad (\forall y \forall y' ((\text{maps}(f, x, y) \wedge \text{maps}(f, x, y')) \rightarrow y = y')))) \end{aligned}$$

where $\text{maps}(f, x, y)$ abbreviates $\exists v (v \in f \wedge \langle x, y \rangle = v)$ (this formula expresses “ $f(x) = y$ ”).

It is now also not hard to express that $f: X \rightarrow Y$ is **injective**, for instance:

$$\begin{aligned} f: X \rightarrow Y \wedge \forall x \forall x' ((x \in X \wedge x' \in X \wedge \\ \exists y (\text{maps}(f, x, y) \wedge \text{maps}(f, x', y))) \rightarrow x = x') \end{aligned}$$

A function $f: X \rightarrow Y$ is **injective** iff, whenever f maps $x, x' \in X$ to a single y , $x = x'$. If we abbreviate this formula as $\text{inj}(f, X, Y)$, we’re already in a position to state in the language of set theory something as non-trivial as Cantor’s theorem: there is no **injective** function from $\wp(X)$ to X :

$$\forall X \forall Y (\wp(X) = Y \rightarrow \neg \exists f \text{inj}(f, Y, X))$$

One might think that set theory requires another axiom that guarantees the existence of a set for every defining property. If $\varphi(x)$ is a formula of set theory with the variable x free, we can consider the **sentence**

$$\exists y \forall x (x \in y \leftrightarrow \varphi(x)).$$

This **sentence** states that there is a set y whose **elements** are all and only those x that satisfy $\varphi(x)$. This schema is called the “comprehension principle.” It looks very useful; unfortunately it is inconsistent. Take $\varphi(x) \equiv \neg x \in x$, then the comprehension principle states

$$\exists y \forall x (x \in y \leftrightarrow x \notin x),$$

i.e., it states the existence of a set of all sets that are not **elements** of themselves. No such set can exist—this is Russell’s Paradox. **ZFC**, in fact, contains a restricted—and consistent—version of this principle, the separation principle:

$$\forall z \exists y \forall x (x \in y \leftrightarrow (x \in z \wedge \varphi(x))).$$

Problem 2.4. Show that the comprehension principle is inconsistent by giving a **derivation** that shows

$$\exists y \forall x (x \in y \leftrightarrow x \notin x) \vdash \perp.$$

It may help to first show $(A \rightarrow \neg A) \wedge (\neg A \rightarrow A) \vdash \perp$.

2.6 Expressing the Size of Structures

explanation

There are some properties of structures we can express even without using the non-logical symbols of a language. For instance, there are **sentences** which are true in a **structure** iff the **domain** of the **structure** has at least, at most, or exactly a certain number n of **elements**.

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Proposition 2.11. *The sentence*

$$\begin{aligned} \varphi_{\geq n} \equiv \exists x_1 \exists x_2 \dots \exists x_n \quad & (x_1 \neq x_2 \wedge x_1 \neq x_3 \wedge x_1 \neq x_4 \wedge \dots \wedge x_1 \neq x_n \wedge \\ & x_2 \neq x_3 \wedge x_2 \neq x_4 \wedge \dots \wedge x_2 \neq x_n \wedge \\ & \vdots \\ & x_{n-1} \neq x_n) \end{aligned}$$

is true in a **structure** \mathfrak{M} iff $|\mathfrak{M}|$ contains at least n **elements**. Consequently, $\mathfrak{M} \models \neg \varphi_{\geq n+1}$ iff $|\mathfrak{M}|$ contains at most n **elements**.

Proposition 2.12. *The sentence*

$$\begin{aligned} \varphi_{=n} \equiv \exists x_1 \exists x_2 \dots \exists x_n \quad & (x_1 \neq x_2 \wedge x_1 \neq x_3 \wedge x_1 \neq x_4 \wedge \dots \wedge x_1 \neq x_n \wedge \\ & x_2 \neq x_3 \wedge x_2 \neq x_4 \wedge \dots \wedge x_2 \neq x_n \wedge \\ & \vdots \\ & x_{n-1} \neq x_n \wedge \\ & \forall y (y = x_1 \vee \dots \vee y = x_n) \dots) \end{aligned}$$

is true in a **structure** \mathfrak{M} iff $|\mathfrak{M}|$ contains exactly n **elements**.

Proposition 2.13. *A structure is infinite iff it is a model of*

$$\{\varphi_{\geq 1}, \varphi_{\geq 2}, \varphi_{\geq 3}, \dots\}$$

There is no single purely logical sentence which is true in \mathfrak{M} iff $|\mathfrak{M}|$ is infinite. However, one can give **sentences** with non-logical **predicate symbols** which only have infinite models (although not every infinite **structure** is a model of them). The property of being a finite structure, and the property of being a **non-enumerable** structure cannot even be expressed with an infinite set of **sentences**. These facts follow from the compactness and Löwenheim-Skolem theorems.

Chapter 3

Derivation Systems

3.1 Introduction

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Logics commonly have both a semantics and a **derivation** system. The semantics concerns concepts such as truth, satisfiability, validity, and entailment. The purpose of **derivation** systems is to provide a purely syntactic method of establishing entailment and validity. They are purely syntactic in the sense that a **derivation** in such a system is a finite syntactic object, usually a sequence (or other finite arrangement) of **sentences** or **formulas**. Good **derivation** systems have the property that any given sequence or arrangement of **sentences** or **formulas** can be verified mechanically to be “correct.”

The simplest (and historically first) **derivation** systems for first-order logic were *axiomatic*. A sequence of **formulas** counts as a **derivation** in such a system if each individual **formula** in it is either among a fixed set of “axioms” or follows from **formulas** coming before it in the sequence by one of a fixed number of “inference rules”—and it can be mechanically verified if a **formula** is an axiom and whether it follows correctly from other **formulas** by one of the inference rules. Axiomatic proof systems are easy to describe—and also easy to handle meta-theoretically—but **derivations** in them are hard to read and understand, and are also hard to produce.

Other **derivation** systems have been developed with the aim of making it easier to construct **derivations** or easier to understand **derivations** once they are complete. Examples are natural deduction, truth trees, also known as tableaux proofs, and the sequent calculus. Some **derivation** systems are designed especially with mechanization in mind, e.g., the resolution method is easy to implement in software (but its **derivations** are essentially impossible to understand). Most of these other proof systems represent **derivations** as trees of **formulas** rather than sequences. This makes it easier to see which parts of a **derivation** depend on which other parts.

So for a given logic, such as first-order logic, the different **derivation** systems will give different explications of what it is for a **sentence** to be a *theorem* and what it means for a **sentence** to be **derivable** from some others. However that is

done (via axiomatic **derivations**, natural deductions, sequent **derivations**, truth trees, resolution refutations), we want these relations to match the semantic notions of validity and entailment. Let's write $\vdash \varphi$ for “ φ is a theorem” and “ $\Gamma \vdash \varphi$ ” for “ φ is **derivable** from Γ .” However \vdash is defined, we want it to match up with \models , that is:

1. $\vdash \varphi$ if and only if $\models \varphi$
2. $\Gamma \vdash \varphi$ if and only if $\Gamma \models \varphi$

The “only if” direction of the above is called *soundness*. A **derivation** system is sound if **derivability** guarantees entailment (or validity). Every decent **derivation** system has to be sound; unsound **derivation** systems are not useful at all. After all, the entire purpose of a **derivation** is to provide a syntactic guarantee of validity or entailment. We'll prove soundness for the **derivation** systems we present.

The converse “if” direction is also important: it is called *completeness*. A complete **derivation** system is strong enough to show that φ is a theorem whenever φ is valid, and that there $\Gamma \vdash \varphi$ whenever $\Gamma \models \varphi$. Completeness is harder to establish, and some logics have no complete **derivation** systems. First-order logic does. Kurt Gödel was the first one to prove completeness for a **derivation** system of first-order logic in his 1929 dissertation.

Another concept that is connected to **derivation** systems is that of *consistency*. A set of **sentences** is called inconsistent if anything whatsoever can be **derived** from it, and consistent otherwise. Inconsistency is the syntactic counterpart to unsatisfiability: like unsatisfiable sets, inconsistent sets of **sentences** do not make good theories, they are defective in a fundamental way. Consistent sets of **sentences** may not be true or useful, but at least they pass that minimal threshold of logical usefulness. For different **derivation** systems the specific definition of consistency of sets of **sentences** might differ, but like \vdash , we want consistency to coincide with its semantic counterpart, satisfiability. We want it to always be the case that Γ is consistent if and only if it is satisfiable. Here, the “if” direction amounts to completeness (consistency guarantees satisfiability), and the “only if” direction amounts to soundness (satisfiability guarantees consistency). In fact, for classical first-order logic, the two versions of soundness and completeness are equivalent.

3.2 The Sequent Calculus

While many **derivation** systems operate with arrangements of **sentences**, the sequent calculus operates with *sequents*. A sequent is an expression of the form

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$$\varphi_1, \dots, \varphi_m \Rightarrow \psi_1, \dots, \psi_n,$$

that is a pair of sequences of **sentences**, separated by the sequent symbol \Rightarrow . Either sequence may be empty. A **derivation** in the sequent calculus is a tree of sequents, where the topmost sequents are of a special form (they are called

“initial sequents” or “axioms”) and every other sequent follows from the sequents immediately above it by one of the rules of inference. The rules of inference either manipulate the **sentences** in the sequents (adding, removing, or rearranging them on either the left or the right), or they introduce a complex **formula** in the conclusion of the rule. For instance, the \wedge L rule allows the inference from $\varphi, \Gamma \Rightarrow \Delta$ to $A \wedge \psi, \Gamma \Rightarrow \Delta$, and the \rightarrow R allows the inference from $\varphi, \Gamma \Rightarrow \Delta, \psi$ to $\Gamma \Rightarrow \Delta, \varphi \rightarrow \psi$, for any Γ, Δ, φ , and ψ . (In particular, Γ and Δ may be empty.)

The \vdash relation based on the sequent calculus is defined as follows: $\Gamma \vdash \varphi$ iff there is some sequence Γ_0 such that every φ in Γ_0 is in Γ and there is a **derivation** with the sequent $\Gamma_0 \Rightarrow \varphi$ at its root. φ is a theorem in the sequent calculus if the sequent $\Rightarrow \varphi$ has a **derivation**. For instance, here is a **derivation** that shows that $\vdash (\varphi \wedge \psi) \rightarrow \varphi$:

$$\frac{\frac{\varphi \Rightarrow \varphi}{\varphi \wedge \psi \Rightarrow \varphi} \wedge\text{L}}{\Rightarrow (\varphi \wedge \psi) \rightarrow \varphi} \rightarrow\text{R}$$

A set Γ is inconsistent in the sequent calculus if there is a **derivation** of $\Gamma_0 \Rightarrow$ (where every $\varphi \in \Gamma_0$ is in Γ and the right side of the sequent is empty). Using the rule WR, any **sentence** can be **derived** from an inconsistent set.

The sequent calculus was invented in the 1930 by Gerhard Gentzen. Because of its systematic and symmetric design, it is a very useful formalism for developing a theory of **derivations**. It is relatively easy to find **derivations** in the sequent calculus, but these **derivations** are often hard to read and their connection to proofs are sometimes not easy to see. It has proved to be a very elegant approach to **derivation** systems, however, and many logics have sequent calculus systems.

3.3 Natural Deduction

fol:prf:ntd:
sec

Natural deduction is a **derivation** system intended to mirror actual reasoning (especially the kind of regimented reasoning employed by mathematicians). Actual reasoning proceeds by a number of “natural” patterns. For instance, proof by cases allows us to establish a conclusion on the basis of a disjunctive premise, by establishing that the conclusion follows from either of the disjuncts. Indirect proof allows us to establish a conclusion by showing that its negation leads to a contradiction. Conditional proof establishes a conditional claim “if ... then ...” by showing that the consequent follows from the antecedent. Natural deduction is a formalization of some of these natural inferences. Each of the logical connectives and quantifiers comes with two rules, an introduction and an elimination rule, and they each correspond to one such natural inference pattern. For instance, \rightarrow Intro corresponds to conditional proof, and \vee Elim to proof by cases. A particularly simple rule is \wedge Elim which allows the inference from $\varphi \wedge \psi$ to φ (or ψ).

One feature that distinguishes natural deduction from other **derivation** systems is its use of assumptions. A **derivation** in natural deduction is a tree of **formulas**. A single **formula** stands at the root of the tree of **formulas**, and the “leaves” of the tree are **formulas** from which the conclusion is derived. In natural deduction, some leaf **formulas** play a role inside the **derivation** but are “used up” by the time the **derivation** reaches the conclusion. This corresponds to the practice, in actual reasoning, of introducing hypotheses which only remain in effect for a short while. For instance, in a proof by cases, we assume the truth of each of the disjuncts; in conditional proof, we assume the truth of the antecedent; in indirect proof, we assume the truth of the negation of the conclusion. This way of introducing hypothetical assumptions and then doing away with them in the service of establishing an intermediate step is a hallmark of natural deduction. The formulas at the leaves of a natural deduction **derivation** are called assumptions, and some of the rules of inference may “**discharge**” them. For instance, if we have a **derivation** of ψ from some assumptions which include φ , then the \rightarrow Intro rule allows us to infer $\varphi \rightarrow \psi$ and discharge any assumption of the form φ . (To keep track of which assumptions are discharged at which inferences, we label the inference and the assumptions it discharges with a number.) The assumptions that remain **undischarged** at the end of the **derivation** are together sufficient for the truth of the conclusion, and so a **derivation** establishes that its **undischarged** assumptions entail its conclusion.

The relation $\Gamma \vdash \varphi$ based on natural deduction holds iff there is a **derivation** in which φ is the last **sentence** in the tree, and every leaf which is **undischarged** is in Γ . φ is a theorem in natural deduction iff there is a **derivation** in which φ is the last **sentence** and all assumptions are **discharged**. For instance, here is a **derivation** that shows that $\vdash (\varphi \wedge \psi) \rightarrow \varphi$:

$$1 \frac{\frac{[\varphi \wedge \psi]^1}{\varphi} \wedge\text{Elim}}{(\varphi \wedge \psi) \rightarrow \varphi} \rightarrow\text{Intro}$$

The label 1 indicates that the assumption $\varphi \wedge \psi$ is **discharged** at the \rightarrow Intro inference.

A set Γ is inconsistent iff $\Gamma \vdash \perp$ in natural deduction. The rule \perp_I makes it so that from an inconsistent set, any **sentence** can be **derived**.

Natural deduction systems were developed by Gerhard Gentzen and Stanisław Jaśkowski in the 1930s, and later developed by Dag Prawitz and Frederic Fitch. Because its inferences mirror natural methods of proof, it is favored by philosophers. The versions developed by Fitch are often used in introductory logic textbooks. In the philosophy of logic, the rules of natural deduction have sometimes been taken to give the meanings of the logical operators (“proof-theoretic semantics”).

3.4 Axiomatic Derivations

fol:prf:axd:
sec

Axiomatic **derivations** are the oldest and simplest logical **derivation** systems. Its **derivations** are simply sequences of **sentences**. A sequence of **sentences** counts as a correct **derivation** if every **sentence** φ in it satisfies one of the following conditions:

1. φ is an axiom, or
2. φ is an **element** of a given set Γ of **sentences**, or
3. φ is justified by a rule of inference.

To be an axiom, φ has to have the form of one of a number of fixed **sentence** schemas. There are many sets of axiom schemas that provide a satisfactory (sound and complete) **derivation** system for first-order logic. Some are organized according to the connectives they govern, e.g., the schemas

$$\varphi \rightarrow (\psi \rightarrow \varphi) \quad \psi \rightarrow (\psi \vee \chi) \quad (\psi \wedge \chi) \rightarrow \psi$$

are common axioms that govern \rightarrow , \vee and \wedge . Some axiom systems aim at a minimal number of axioms. Depending on the connectives that are taken as primitives, it is even possible to find axiom systems that consist of a single axiom.

A rule of inference is a conditional statement that gives a sufficient condition for a **sentence** in a **derivation** to be justified. Modus ponens is one very common such rule: it says that if φ and $\varphi \rightarrow \psi$ are already justified, then ψ is justified. This means that a line in a **derivation** containing the **sentence** ψ is justified, provided that both φ and $\varphi \rightarrow \psi$ (for some **sentence** φ) appear in the **derivation** before ψ .

The \vdash relation based on axiomatic **derivations** is defined as follows: $\Gamma \vdash \varphi$ iff there is a **derivation** with the **sentence** φ as its last formula (and Γ is taken as the set of **sentences** in that derivation which are justified by (2) above). φ is a theorem if φ has a **derivation** where Γ is empty, i.e., every **sentence** in the derivation is justified either by (1) or (3). For instance, here is a **derivation** that shows that $\vdash \varphi \rightarrow (\psi \rightarrow (\psi \vee \varphi))$:

1. $\psi \rightarrow (\psi \vee \varphi)$
2. $(\psi \rightarrow (\psi \vee \varphi)) \rightarrow (\varphi \rightarrow (\psi \rightarrow (\psi \vee \varphi)))$
3. $\varphi \rightarrow (\psi \rightarrow (\psi \vee \varphi))$

The **sentence** on line 1 is of the form of the axiom $\varphi \rightarrow (\varphi \vee \psi)$ (with the roles of φ and ψ reversed). The sentence on line 2 is of the form of the axiom $\varphi \rightarrow (\psi \rightarrow \varphi)$. Thus, both lines are justified. Line 3 is justified by modus ponens: if we abbreviate it as θ , then line 2 has the form $\chi \rightarrow \theta$, where χ is $\psi \rightarrow (\psi \vee \varphi)$, i.e., line 1.

A set Γ is inconsistent if $\Gamma \vdash \perp$. A complete axiom system will also prove that $\perp \rightarrow \varphi$ for any φ , and so if Γ is inconsistent, then $\Gamma \vdash \varphi$ for any φ .

Systems of axiomatic **derivations** for logic were first given by Gottlob Frege in his 1879 *Begriffsschrift*, which for this reason is often considered the first work of modern logic. They were perfected in Alfred North Whitehead and Bertrand Russell's *Principia Mathematica* and by David Hilbert and his students in the 1920s. They are thus often called "Frege systems" or "Hilbert systems." They are very versatile in that it is often easy to find an axiomatic system for a logic. Because **derivations** have a very simple structure and only one or two inference rules, it is also relatively easy to prove things *about* them. However, they are very hard to use in practice, i.e., it is difficult to find and write proofs.

Chapter 4

The Sequent Calculus

4.1 Rules and Derivations

fol:seq:rul:
sec For the following, let $\Gamma, \Delta, \Pi, \Lambda$ represent finite sequences of **sentences**.

Definition 4.1 (Sequent). A *sequent* is an expression of the form

$$\Gamma \Rightarrow \Delta$$

where Γ and Δ are finite (possibly empty) sequences of **sentences** of the language \mathcal{L} . Γ is called the *antecedent*, while Δ is the *succedent*.

The intuitive idea behind a sequent is: if all of the **sentences** in the antecedent hold, then at least one of the **sentences** in the succedent holds. That is, if $\Gamma = \langle \varphi_1, \dots, \varphi_m \rangle$ and $\Delta = \langle \psi_1, \dots, \psi_n \rangle$, then $\Gamma \Rightarrow \Delta$ holds iff explanation

$$(\varphi_1 \wedge \dots \wedge \varphi_m) \rightarrow (\psi_1 \vee \dots \vee \psi_n)$$

holds. There are two special cases: where Γ is empty and when Δ is empty. When Γ is empty, i.e., $m = 0$, $\Rightarrow \Delta$ holds iff $\psi_1 \vee \dots \vee \psi_n$ holds. When Δ is empty, i.e., $n = 0$, $\Gamma \Rightarrow$ holds iff $\neg(\varphi_1 \wedge \dots \wedge \varphi_m)$ does. We say a sequent is valid iff the corresponding **sentence** is valid.

If Γ is a sequence of **sentences**, we write Γ, φ for the result of appending φ to the right end of Γ (and φ, Γ for the result of appending φ to the left end of Γ). If Δ is a sequence of **sentences** also, then Γ, Δ is the concatenation of the two sequences.

Definition 4.2 (Initial Sequent). An *initial sequent* is a sequent of one of the following forms:

1. $\varphi \Rightarrow \varphi$
2. $\Rightarrow \top$
3. $\perp \Rightarrow$

for any **sentence** φ in the language.

Derivations in the sequent calculus are certain trees of sequents, where the topmost sequents are initial sequents, and if a sequent stands below one or two other sequents, it must follow correctly by a rule of inference. The rules for **LK** are divided into two main types: *logical* rules and *structural* rules. The logical rules are named for the **main operator** of the **sentence** containing φ and/or ψ in the lower sequent. Each one comes in two versions, one for inferring a sequent with the **sentence** containing the **logical operator** on the left, and one with the **sentence** on the right.

4.2 Propositional Rules

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sec

Rules for \neg

$$\frac{\Gamma \Rightarrow \Delta, \varphi}{\neg\varphi, \Gamma \Rightarrow \Delta} \neg\text{L} \qquad \frac{\varphi, \Gamma \Rightarrow \Delta}{\Gamma \Rightarrow \Delta, \neg\varphi} \neg\text{R}$$

Rules for \wedge

$$\frac{\varphi, \Gamma \Rightarrow \Delta}{\varphi \wedge \psi, \Gamma \Rightarrow \Delta} \wedge\text{L} \qquad \frac{\Gamma \Rightarrow \Delta, \varphi \quad \Gamma \Rightarrow \Delta, \psi}{\Gamma \Rightarrow \Delta, \varphi \wedge \psi} \wedge\text{R}$$

$$\frac{\psi, \Gamma \Rightarrow \Delta}{\varphi \wedge \psi, \Gamma \Rightarrow \Delta} \wedge\text{L}$$

Rules for \vee

$$\frac{\varphi, \Gamma \Rightarrow \Delta \quad \psi, \Gamma \Rightarrow \Delta}{\varphi \vee \psi, \Gamma \Rightarrow \Delta} \vee\text{L} \qquad \frac{\Gamma \Rightarrow \Delta, \varphi}{\Gamma \Rightarrow \Delta, \varphi \vee \psi} \vee\text{R}$$

$$\frac{\Gamma \Rightarrow \Delta, \psi}{\Gamma \Rightarrow \Delta, \varphi \vee \psi} \vee\text{R}$$

Rules for \rightarrow

$$\frac{\Gamma \Rightarrow \Delta, \varphi \quad \psi, \Pi \Rightarrow \Lambda}{\varphi \rightarrow \psi, \Gamma, \Pi \Rightarrow \Delta, \Lambda} \rightarrow\text{L} \qquad \frac{\varphi, \Gamma \Rightarrow \Delta, \psi}{\Gamma \Rightarrow \Delta, \varphi \rightarrow \psi} \rightarrow\text{R}$$

4.3 Quantifier Rules

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sec

Rules for \forall

$$\boxed{\frac{\varphi(t), \Gamma \Rightarrow \Delta}{\forall x \varphi(x), \Gamma \Rightarrow \Delta} \forall L \qquad \frac{\Gamma \Rightarrow \Delta, \varphi(a)}{\Gamma \Rightarrow \Delta, \forall x \varphi(x)} \forall R}$$

In $\forall L$, t is a closed term (i.e., one without variables). In $\forall R$, a is a **constant symbol** which must not occur anywhere in the lower sequent of the $\forall R$ rule. We call a the *eigenvariable* of the $\forall R$ inference.

Rules for \exists

$$\boxed{\frac{\varphi(a), \Gamma \Rightarrow \Delta}{\exists x \varphi(x), \Gamma \Rightarrow \Delta} \exists L \qquad \frac{\Gamma \Rightarrow \Delta, \varphi(t)}{\Gamma \Rightarrow \Delta, \exists x \varphi(x)} \exists R}$$

Again, t is a closed term, and a is a **constant symbol** which does not occur in the lower sequent of the $\exists L$ rule. We call a the *eigenvariable* of the $\exists L$ inference.

The condition that an eigenvariable not occur in the lower sequent of the $\forall R$ or $\exists L$ inference is called the *eigenvariable condition*.

We use the term “eigenvariable” even though a in the above rules is a **constant symbol**. This has historical reasons. explanation

In $\exists R$ and $\forall L$ there are no restrictions on the term t . On the other hand, in the $\exists L$ and $\forall R$ rules, the eigenvariable condition requires that the **constant symbol** a does not occur anywhere outside of $\varphi(a)$ in the upper sequent. It is necessary to ensure that the system is sound, i.e., only **derives** sequents that are valid. Without this condition, the following would be allowed:

$$\frac{\frac{\varphi(a) \Rightarrow \varphi(a)}{\exists x \varphi(x) \Rightarrow \varphi(a)} * \exists L}{\exists x \varphi(x) \Rightarrow \forall x \varphi(x)} \forall R \qquad \frac{\frac{\varphi(a) \Rightarrow \varphi(a)}{\varphi(a) \Rightarrow \forall x \varphi(x)} * \forall R}{\exists x \varphi(x) \Rightarrow \forall x \varphi(x)} \exists L$$

However, $\exists x \varphi(x) \Rightarrow \forall x \varphi(x)$ is not valid.

4.4 Structural Rules

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sec

We also need a few rules that allow us to rearrange **sentences** in the left and right side of a sequent. Since the logical rules require that the **sentences** in

the premise which the rule acts upon stand either to the far left or to the far right, we need an “exchange” rule that allows us to move **sentences** to the right position. It’s also important sometimes to be able to combine two identical **sentences** into one, and to add a **sentence** on either side.

Weakening

$$\frac{\Gamma \Rightarrow \Delta}{\varphi, \Gamma \Rightarrow \Delta} \text{WL} \qquad \frac{\Gamma \Rightarrow \Delta}{\Gamma \Rightarrow \Delta, \varphi} \text{WR}$$

Contraction

$$\frac{\varphi, \varphi, \Gamma \Rightarrow \Delta}{\varphi, \Gamma \Rightarrow \Delta} \text{CL} \qquad \frac{\Gamma \Rightarrow \Delta, \varphi, \varphi}{\Gamma \Rightarrow \Delta, \varphi} \text{CR}$$

Exchange

$$\frac{\Gamma, \varphi, \psi, \Pi \Rightarrow \Delta}{\Gamma, \psi, \varphi, \Pi \Rightarrow \Delta} \text{XL} \qquad \frac{\Gamma \Rightarrow \Delta, \varphi, \psi, \Lambda}{\Gamma \Rightarrow \Delta, \psi, \varphi, \Lambda} \text{XR}$$

A series of weakening, contraction, and exchange inferences will often be indicated by double inference lines.

The following rule, called “cut,” is not strictly speaking necessary, but makes it a lot easier to reuse and combine derivations.

$$\frac{\Gamma \Rightarrow \Delta, \varphi \quad \varphi, \Pi \Rightarrow \Lambda}{\Gamma, \Pi \Rightarrow \Delta, \Lambda} \text{Cut}$$

4.5 Derivations

explanation We’ve said what an initial sequent looks like, and we’ve given the rules of inference. **Derivations** in the sequent calculus are inductively generated from these: each **derivation** either is an initial sequent on its own, or consists of one or two **derivations** followed by an inference. fol:seq:der:sec

Definition 4.3 (LK derivation). An **LK-derivation** of a sequent S is a tree of sequents satisfying the following conditions:

1. The topmost sequents of the tree are initial sequents.

2. The bottommost sequent of the tree is S .
3. Every sequent in the tree except S is a premise of a correct application of an inference rule whose conclusion stands directly below that sequent in the tree.

We then say that S is the *end-sequent* of the *derivation* and that S is *derivable in LK* (or *LK-derivable*).

Example 4.4. Every initial sequent, e.g., $\chi \Rightarrow \chi$ is a *derivation*. We can obtain a new *derivation* from this by applying, say, the WL rule,

$$\frac{\Gamma \Rightarrow \Delta}{\varphi, \Gamma \Rightarrow \Delta} \text{WL}$$

The rule, however, is meant to be general: we can replace the φ in the rule with any *sentence*, e.g., also with θ . If the premise matches our initial sequent $\chi \Rightarrow \chi$, that means that both Γ and Δ are just χ , and the conclusion would then be $\theta, \chi \Rightarrow \chi$. So, the following is a *derivation*:

$$\frac{\chi \Rightarrow \chi}{\theta, \chi \Rightarrow \chi} \text{WL}$$

We can now apply another rule, say XL, which allows us to switch two *sentences* on the left. So, the following is also a correct *derivation*:

$$\frac{\frac{\chi \Rightarrow \chi}{\theta, \chi \Rightarrow \chi} \text{WL}}{\chi, \theta \Rightarrow \chi} \text{XL}$$

In this application of the rule, which was given as

$$\frac{\Gamma, \varphi, \psi, \Pi \Rightarrow \Delta}{\Gamma, \psi, \varphi, \Pi \Rightarrow \Delta} \text{XL}$$

both Γ and Π were empty, Δ is χ , and the roles of φ and ψ are played by θ and χ , respectively. In much the same way, we also see that

$$\frac{\theta \Rightarrow \theta}{\chi, \theta \Rightarrow \theta} \text{WL}$$

is a *derivation*. Now we can take these two derivations, and combine them using $\wedge\text{R}$. That rule was

$$\frac{\Gamma \Rightarrow \Delta, \varphi \quad \Gamma \Rightarrow \Delta, \psi}{\Gamma \Rightarrow \Delta, \varphi \wedge \psi} \wedge\text{R}$$

In our case, the premises must match the last sequents of the *derivations* ending in the premises. That means that Γ is χ, θ , Δ is empty, φ is χ and ψ is θ . So the conclusion, if the inference should be correct, is $\chi, \theta \Rightarrow \chi \wedge \theta$. Of course, we can also reverse the premises, then φ would be θ and ψ would be χ . So both of the following are correct *derivations*.

$$\frac{\frac{\frac{\chi \Rightarrow \chi}{\theta, \chi \Rightarrow \chi} \text{WL}}{\chi, \theta \Rightarrow \chi} \text{XL}}{\chi, \theta \Rightarrow \chi \wedge \theta} \wedge \text{R} \quad \frac{\frac{\theta \Rightarrow \theta}{\chi, \theta \Rightarrow \theta} \text{WL}}{\chi, \theta \Rightarrow \theta \wedge \chi} \wedge \text{R}$$

4.6 Examples of Derivations

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sec

Example 4.5. Give an **LK**-derivation for the sequent $\varphi \wedge \psi \Rightarrow \varphi$.

We begin by writing the desired end-sequent at the bottom of the derivation.

$$\overline{\varphi \wedge \psi \Rightarrow \varphi}$$

Next, we need to figure out what kind of inference could have a lower sequent of this form. This could be a structural rule, but it is a good idea to start by looking for a logical rule. The only logical connective occurring in the lower sequent is \wedge , so we're looking for an \wedge rule, and since the \wedge symbol occurs in the antecedent, we're looking at the $\wedge \text{L}$ rule.

$$\overline{\varphi \wedge \psi \Rightarrow \varphi} \wedge \text{L}$$

There are two options for what could have been the upper sequent of the $\wedge \text{L}$ inference: we could have an upper sequent of $\varphi \Rightarrow \varphi$, or of $\psi \Rightarrow \varphi$. Clearly, $\varphi \Rightarrow \varphi$ is an initial sequent (which is a good thing), while $\psi \Rightarrow \varphi$ is not derivable in general. We fill in the upper sequent:

$$\frac{\varphi \Rightarrow \varphi}{\varphi \wedge \psi \Rightarrow \varphi} \wedge \text{L}$$

We now have a correct **LK**-derivation of the sequent $\varphi \wedge \psi \Rightarrow \varphi$.

Example 4.6. Give an **LK**-derivation for the sequent $\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi$.

Begin by writing the desired end-sequent at the bottom of the derivation.

$$\overline{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi}$$

To find a logical rule that could give us this end-sequent, we look at the logical connectives in the end-sequent: \neg , \vee , and \rightarrow . We only care at the moment about \vee and \rightarrow because they are **main operators of sentences** in the end-sequent, while \neg is inside the scope of another connective, so we will take care of it later. Our options for logical rules for the final inference are therefore the $\vee \text{L}$ rule and the $\rightarrow \text{R}$ rule. We could pick either rule, really, but let's pick the $\rightarrow \text{R}$ rule (if for no reason other than it allows us to put off splitting into two branches). According to the form of $\rightarrow \text{R}$ inferences which can yield the lower sequent, this must look like:

$$\frac{\overline{\varphi, \neg\varphi \vee \psi \Rightarrow \psi}}{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi} \rightarrow \text{R}$$

If we move $\neg\varphi \vee \psi$ to the outside of the antecedent, we can apply the $\vee\text{L}$ rule. According to the schema, this must split into two upper sequents as follows:

$$\frac{\frac{\frac{\overline{\neg\varphi, \varphi \Rightarrow \psi}}{\neg\varphi \vee \psi, \varphi \Rightarrow \psi} \vee\text{L}}{\varphi, \neg\varphi \vee \psi \Rightarrow \psi} \text{XR}}{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi} \rightarrow\text{R}$$

Remember that we are trying to wind our way up to initial sequents; we seem to be pretty close! The right branch is just one weakening and one exchange away from an initial sequent and then it is done:

$$\frac{\frac{\frac{\overline{\neg\varphi, \varphi \Rightarrow \psi}}{\neg\varphi \vee \psi, \varphi \Rightarrow \psi} \vee\text{L}}{\varphi, \neg\varphi \vee \psi \Rightarrow \psi} \text{XR}}{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi} \rightarrow\text{R}$$

Now looking at the left branch, the only logical connective in any **sentence** is the \neg symbol in the antecedent **sentences**, so we're looking at an instance of the $\neg\text{L}$ rule.

$$\frac{\frac{\frac{\overline{\varphi \Rightarrow \psi, \varphi}}{\neg\varphi, \varphi \Rightarrow \psi} \neg\text{L}}{\neg\varphi \vee \psi, \varphi \Rightarrow \psi} \vee\text{L}}{\varphi, \neg\varphi \vee \psi \Rightarrow \psi} \text{XR}}{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi} \rightarrow\text{R}$$

Similarly to how we finished off the right branch, we are just one weakening and one exchange away from finishing off this left branch as well.

$$\frac{\frac{\frac{\frac{\overline{\varphi \Rightarrow \varphi}}{\varphi \Rightarrow \varphi, \psi} \text{WR}}{\varphi \Rightarrow \psi, \varphi} \text{XR}}{\neg\varphi, \varphi \Rightarrow \psi} \neg\text{L}}{\neg\varphi \vee \psi, \varphi \Rightarrow \psi} \vee\text{L}}{\varphi, \neg\varphi \vee \psi \Rightarrow \psi} \text{XR}}{\neg\varphi \vee \psi \Rightarrow \varphi \rightarrow \psi} \rightarrow\text{R}$$

Example 4.7. Give an **LK**-derivation of the sequent $\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)$

Using the techniques from above, we start by writing the desired end-sequent at the bottom.

$$\overline{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)}$$

The available main connectives of **sentences** in the end-sequent are the \vee symbol and the \neg symbol. It would work to apply either the $\vee\text{L}$ or the $\neg\text{R}$ rule here, but we start with the $\neg\text{R}$ rule because it avoids splitting up into two branches for a moment:

$$\frac{\overline{\varphi \wedge \psi, \neg\varphi \vee \neg\psi \Rightarrow}}{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)} \neg\text{R}$$

Now we have a choice of whether to look at the $\wedge\text{L}$ or the $\vee\text{L}$ rule. Let's see what happens when we apply the $\wedge\text{L}$ rule: we have a choice to start with either the sequent $\varphi, \neg\varphi \vee \neg\psi \Rightarrow$ or the sequent $\psi, \neg\varphi \vee \neg\psi \Rightarrow$. Since the proof is symmetric with regards to φ and ψ , let's go with the former:

$$\frac{\frac{\overline{\varphi, \neg\varphi \vee \neg\psi \Rightarrow}}{\varphi \wedge \psi, \neg\varphi \vee \neg\psi \Rightarrow} \wedge\text{L}}{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)} \neg\text{R}$$

Continuing to fill in the derivation, we see that we run into a problem:

$$\frac{\frac{\frac{\frac{\overline{\varphi \Rightarrow \varphi}}{\neg\varphi, \varphi \Rightarrow} \neg\text{L} \quad \frac{\overline{\varphi \Rightarrow \psi} \quad ?}{\neg\psi, \varphi \Rightarrow} \neg\text{L}}{\neg\varphi \vee \neg\psi, \varphi \Rightarrow} \vee\text{L}}{\frac{\overline{\varphi, \neg\varphi \vee \neg\psi \Rightarrow} \text{XL}}{\varphi \wedge \psi, \neg\varphi \vee \neg\psi \Rightarrow} \wedge\text{L}}{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)} \neg\text{R}$$

The top of the right branch cannot be reduced any further, and it cannot be brought by way of structural inferences to an initial sequent, so this is not the right path to take. So clearly, it was a mistake to apply the $\wedge\text{L}$ rule above. Going back to what we had before and carrying out the $\vee\text{L}$ rule instead, we get

$$\frac{\frac{\overline{\neg\varphi, \varphi \wedge \psi \Rightarrow} \quad \overline{\neg\psi, \varphi \wedge \psi \Rightarrow}}{\neg\varphi \vee \neg\psi, \varphi \wedge \psi \Rightarrow} \vee\text{L}}{\frac{\overline{\varphi \wedge \psi, \neg\varphi \vee \neg\psi \Rightarrow} \text{XL}}{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)} \neg\text{R}}$$

Completing each branch as we've done before, we get

$$\frac{\frac{\frac{\overline{\varphi \Rightarrow \varphi}}{\varphi \wedge \psi \Rightarrow \varphi} \wedge\text{L} \quad \frac{\overline{\psi \Rightarrow \psi}}{\varphi \wedge \psi \Rightarrow \psi} \wedge\text{L}}{\neg\varphi, \varphi \wedge \psi \Rightarrow} \neg\text{L} \quad \frac{\overline{\neg\psi, \varphi \wedge \psi \Rightarrow} \neg\text{L}}{\neg\psi, \varphi \wedge \psi \Rightarrow} \neg\text{L}}{\neg\varphi \vee \neg\psi, \varphi \wedge \psi \Rightarrow} \vee\text{L}}{\frac{\overline{\varphi \wedge \psi, \neg\varphi \vee \neg\psi \Rightarrow} \text{XL}}{\neg\varphi \vee \neg\psi \Rightarrow \neg(\varphi \wedge \psi)} \neg\text{R}}$$

(We could have carried out the \wedge rules lower than the \neg rules in these steps and still obtained a correct derivation).

Example 4.8. So far we haven't used the contraction rule, but it is sometimes required. Here's an example where that happens. Suppose we want to prove $\Rightarrow A \vee \neg\varphi$. Applying $\vee\text{R}$ backwards would give us one of these two **derivations**:

$$\frac{\overline{\Rightarrow \varphi}}{\Rightarrow \varphi \vee \neg\varphi} \vee\text{R} \qquad \frac{\overline{\varphi \Rightarrow}}{\Rightarrow \neg\varphi} \neg\text{R} \qquad \frac{\overline{\varphi \Rightarrow}}{\Rightarrow \varphi \vee \neg\varphi} \vee\text{R}$$

Neither of these of course ends in an initial sequent. The trick is to realize that the contraction rule allows us to combine two copies of a **sentence** into one—and when we're searching for a proof, i.e., going from bottom to top, we can keep a copy of $\varphi \vee \neg\varphi$ in the premise, e.g.,

$$\frac{\overline{\Rightarrow \varphi \vee \neg\varphi, \varphi}}{\Rightarrow \varphi \vee \neg\varphi, \varphi \vee \neg\varphi} \vee\text{R} \qquad \frac{\overline{\Rightarrow \varphi \vee \neg\varphi, \varphi \vee \neg\varphi}}{\Rightarrow \varphi \vee \neg\varphi} \text{CR}$$

Now we can apply $\vee\text{R}$ a second time, and also get $\neg\varphi$, which leads to a complete **derivation**.

$$\frac{\overline{\varphi \Rightarrow \varphi}}{\Rightarrow \varphi, \neg\varphi} \neg\text{R} \qquad \frac{\overline{\Rightarrow \varphi, \neg\varphi}}{\Rightarrow \varphi, \varphi \vee \neg\varphi} \vee\text{R} \qquad \frac{\overline{\Rightarrow \varphi, \varphi \vee \neg\varphi}}{\Rightarrow \varphi \vee \neg\varphi, \varphi} \text{XR} \qquad \frac{\overline{\Rightarrow \varphi \vee \neg\varphi, \varphi \vee \neg\varphi}}{\Rightarrow \varphi \vee \neg\varphi} \vee\text{R} \qquad \frac{\overline{\Rightarrow \varphi \vee \neg\varphi, \varphi \vee \neg\varphi}}{\Rightarrow \varphi \vee \neg\varphi} \text{CR}$$

Problem 4.1. Give **derivations** of the following sequents:

1. $\Rightarrow \neg(\varphi \rightarrow \psi) \rightarrow (\varphi \wedge \neg\psi)$
2. $(\varphi \wedge \psi) \rightarrow \chi \Rightarrow (\varphi \rightarrow \chi) \vee (\psi \rightarrow \chi)$

4.7 Derivations with Quantifiers

fol:seq:prq:
sec

Example 4.9. Give an **LK**-derivation of the sequent $\exists x \neg\varphi(x) \Rightarrow \neg\forall x \varphi(x)$.

When dealing with quantifiers, we have to make sure not to violate the eigenvariable condition, and sometimes this requires us to play around with the order of carrying out certain inferences. In general, it helps to try and take care of rules subject to the eigenvariable condition first (they will be lower down in the finished proof). Also, it is a good idea to try and look ahead and try to guess what the initial sequent might look like. In our case, it will have to be something like $\varphi(a) \Rightarrow \varphi(a)$. That means that when we are “reversing” the quantifier rules, we will have to pick the same term—what we will call a —for

both the \forall and the \exists rule. If we picked different terms for each rule, we would end up with something like $\varphi(a) \Rightarrow \varphi(b)$, which, of course, is not derivable.

Starting as usual, we write

$$\overline{\exists x \neg \varphi(x) \Rightarrow \neg \forall x \varphi(x)}$$

We could either carry out the $\exists\text{L}$ rule or the $\neg\text{R}$ rule. Since the $\exists\text{L}$ rule is subject to the eigenvariable condition, it's a good idea to take care of it sooner rather than later, so we'll do that one first.

$$\frac{\overline{\neg \varphi(a) \Rightarrow \neg \forall x \varphi(x)}}{\exists x \neg \varphi(x) \Rightarrow \neg \forall x \varphi(x)} \exists\text{L}$$

Applying the $\neg\text{L}$ and $\neg\text{R}$ rules backwards, we get

$$\frac{\frac{\frac{\overline{\forall x \varphi(x) \Rightarrow \varphi(a)}}{\neg \varphi(a), \forall x \varphi(x) \Rightarrow} \neg\text{L}}{\forall x \varphi(x), \neg \varphi(a) \Rightarrow} \text{XL}}{\neg \varphi(a) \Rightarrow \neg \forall x \varphi(x)} \neg\text{R}}{\exists x \neg \varphi(x) \Rightarrow \neg \forall x \varphi(x)} \exists\text{L}$$

At this point, our only option is to carry out the $\forall\text{L}$ rule. Since this rule is not subject to the eigenvariable restriction, we're in the clear. Remember, we want to try and obtain an initial sequent (of the form $\varphi(a) \Rightarrow \varphi(a)$), so we should choose a as our argument for φ when we apply the rule.

$$\frac{\frac{\frac{\frac{\overline{\varphi(a) \Rightarrow \varphi(a)}}{\forall x \varphi(x) \Rightarrow \varphi(a)} \forall\text{L}}{\neg \varphi(a), \forall x \varphi(x) \Rightarrow} \neg\text{L}}{\forall x \varphi(x), \neg \varphi(a) \Rightarrow} \text{XL}}{\neg \varphi(a) \Rightarrow \neg \forall x \varphi(x)} \neg\text{R}}{\exists x \neg \varphi(x) \Rightarrow \neg \forall x \varphi(x)} \exists\text{L}$$

It is important, especially when dealing with quantifiers, to double check at this point that the eigenvariable condition has not been violated. Since the only rule we applied that is subject to the eigenvariable condition was $\exists\text{L}$, and the eigenvariable a does not occur in its lower sequent (the end-sequent), this is a correct derivation.

Problem 4.2. Give **derivations** of the following sequents:

1. $\forall x (\varphi(x) \rightarrow \psi) \Rightarrow (\exists y \varphi(y) \rightarrow \psi)$
2. $\exists x (\varphi(x) \rightarrow \forall y \varphi(y))$

4.8 Proof-Theoretic Notions

fol:seq:ptn: Just as we've defined a number of important semantic notions (validity, explanation
sec entailment, satisfiability), we now define corresponding *proof-theoretic notions*. These are not defined by appeal to satisfaction of sentences in structures, but by appeal to the derivability or non-derivability of certain sequents. It was an important discovery, due to Gödel, that these notions coincide. That they do is the content of the *completeness theorem*.

Definition 4.10 (Theorems). A sentence φ is a *theorem* if there is a derivation in **LK** of the sequent $\Rightarrow \varphi$. We write $\vdash \varphi$ if φ is a theorem and $\nvdash \varphi$ if it is not.

Definition 4.11 (Derivability). A sentence φ is *derivable* from a set of sentences Γ , $\Gamma \vdash \varphi$, iff there is a finite subset $\Gamma_0 \subseteq \Gamma$ and a sequence Γ'_0 of the sentences in Γ_0 such that **LK** derives $\Gamma'_0 \Rightarrow \varphi$. If φ is not derivable from Γ we write $\Gamma \nvdash \varphi$.

Because of the contraction, weakening, and exchange rules, the order and number of sentences in Γ'_0 does not matter: if a sequent $\Gamma'_0 \Rightarrow \varphi$ is derivable, then so is $\Gamma''_0 \Rightarrow \varphi$ for any Γ''_0 that contains the same sentences as Γ'_0 . For instance, if $\Gamma_0 = \{\psi, \chi\}$ then both $\Gamma'_0 = \langle \psi, \psi, \chi \rangle$ and $\Gamma''_0 = \langle \chi, \chi, \psi \rangle$ are sequences containing just the sentences in Γ_0 . If a sequent containing one is derivable, so is the other, e.g.:

$$\begin{array}{c} \vdots \\ \vdots \\ \vdots \\ \frac{\psi, \psi, \chi \Rightarrow \varphi}{\psi, \chi \Rightarrow \varphi} \text{CL} \\ \frac{\psi, \chi \Rightarrow \varphi}{\chi, \psi \Rightarrow \varphi} \text{XL} \\ \frac{\chi, \psi \Rightarrow \varphi}{\chi, \chi, \psi \Rightarrow \varphi} \text{WL} \end{array}$$

From now on we'll say that if Γ_0 is a finite set of sentences then $\Gamma_0 \Rightarrow \varphi$ is any sequent where the antecedent is a sequence of sentences in Γ_0 and tacitly include contractions, exchanges, and weakenings if necessary.

Definition 4.12 (Consistency). A set of sentences Γ is *inconsistent* iff there is a finite subset $\Gamma_0 \subseteq \Gamma$ such that **LK** derives $\Gamma_0 \Rightarrow$. If Γ is not inconsistent, i.e., if for every finite $\Gamma_0 \subseteq \Gamma$, **LK** does not derive $\Gamma_0 \Rightarrow$, we say it is *consistent*.

fol:seq:ptn: **Proposition 4.13** (Reflexivity). If $\varphi \in \Gamma$, then $\Gamma \vdash \varphi$.
prop:reflexivity

Proof. The initial sequent $\varphi \Rightarrow \varphi$ is derivable, and $\{\varphi\} \subseteq \Gamma$. □

fol:seq:ptn: **Proposition 4.14** (Monotony). If $\Gamma \subseteq \Delta$ and $\Gamma \vdash \varphi$, then $\Delta \vdash \varphi$.
prop:monotony

4.9 Derivability and Consistency

fol:seq:prv:
sec

We will now establish a number of properties of the **derivability** relation. They are independently interesting, but each will play a role in the proof of the completeness theorem.

fol:seq:prv:
prop:provability-contr

Proposition 4.18. *If $\Gamma \vdash \varphi$ and $\Gamma \cup \{\varphi\}$ is inconsistent, then Γ is inconsistent.*

Proof. There are finite Γ_0 and $\Gamma_1 \subseteq \Gamma$ such that **LK** derives $\Gamma_0 \Rightarrow \varphi$ and $\varphi, \Gamma_1 \Rightarrow$. Let the **LK-derivation** of $\Gamma_0 \Rightarrow \varphi$ be π_0 and the **LK-derivation** of $\Gamma_1, \varphi \Rightarrow$ be π_1 . We can then **derive**

$$\frac{\begin{array}{c} \vdots \\ \pi_0 \\ \vdots \\ \Gamma_0 \Rightarrow \varphi \end{array} \quad \begin{array}{c} \vdots \\ \pi_1 \\ \vdots \\ \varphi, \Gamma_1 \Rightarrow \end{array}}{\Gamma_0, \Gamma_1 \Rightarrow} \text{Cut}$$

Since $\Gamma_0 \subseteq \Gamma$ and $\Gamma_1 \subseteq \Gamma$, $\Gamma_0 \cup \Gamma_1 \subseteq \Gamma$, hence Γ is inconsistent. \square

fol:seq:prv:
prop:prov-incons

Proposition 4.19. *$\Gamma \vdash \varphi$ iff $\Gamma \cup \{\neg\varphi\}$ is inconsistent.*

Proof. First suppose $\Gamma \vdash \varphi$, i.e., there is a **derivation** π_0 of $\Gamma \Rightarrow \varphi$. By adding a $\neg\text{L}$ rule, we obtain a **derivation** of $\neg\varphi, \Gamma \Rightarrow$, i.e., $\Gamma \cup \{\neg\varphi\}$ is inconsistent.

If $\Gamma \cup \{\neg A\}$ is inconsistent, there is a **derivation** π_1 of $\neg\varphi, \Gamma \Rightarrow$. The following is a **derivation** of $\Gamma \Rightarrow \varphi$:

$$\frac{\frac{\varphi \Rightarrow \varphi}{\Rightarrow \varphi, \neg\varphi} \neg\text{R} \quad \begin{array}{c} \vdots \\ \pi_1 \\ \vdots \\ \neg\varphi, \Gamma \Rightarrow \end{array}}{\Gamma \Rightarrow \varphi} \text{Cut}$$

\square

Problem 4.4. Prove that $\Gamma \vdash \neg\varphi$ iff $\Gamma \cup \{\varphi\}$ is inconsistent.

fol:seq:prv:
prop:explicit-inc

Proposition 4.20. *If $\Gamma \vdash \varphi$ and $\neg\varphi \in \Gamma$, then Γ is inconsistent.*

Proof. Suppose $\Gamma \vdash \varphi$ and $\neg\varphi \in \Gamma$. Then there is a **derivation** π of a sequent $\Gamma_0 \Rightarrow \varphi$. The sequent $\neg\varphi, \Gamma_0 \Rightarrow$ is also **derivable**:

$$\frac{\begin{array}{c} \vdots \\ \pi \\ \vdots \\ \Gamma_0 \Rightarrow \varphi \end{array} \quad \frac{\frac{\varphi \Rightarrow \varphi}{\neg\varphi, \varphi \Rightarrow} \neg\text{L}}{\varphi, \neg\varphi \Rightarrow} \text{XL}}{\Gamma, \neg\varphi \Rightarrow} \text{Cut}$$

Since $\neg\varphi \in \Gamma$ and $\Gamma_0 \subseteq \Gamma$, this shows that Γ is inconsistent. \square

Proposition 4.21. *If $\Gamma \cup \{\varphi\}$ and $\Gamma \cup \{\neg\varphi\}$ are both inconsistent, then Γ is inconsistent.* fol:seq:prv:
prop:provability-exhaustive

Proof. There are finite sets $\Gamma_0 \subseteq \Gamma$ and $\Gamma_1 \subseteq \Gamma$ and **LK-derivations** π_0 and π_1 of $\varphi, \Gamma_0 \Rightarrow$ and $\neg\varphi, \Gamma_1 \Rightarrow$, respectively. We can then **derive**

$$\frac{\frac{\frac{\vdots \pi_0}{\varphi, \Gamma_0 \Rightarrow}}{\Gamma_0 \Rightarrow \neg\varphi} \neg\text{R} \quad \frac{\vdots \pi_1}{\neg\varphi, \Gamma_1 \Rightarrow}}{\Gamma_0, \Gamma_1 \Rightarrow} \text{Cut}$$

Since $\Gamma_0 \subseteq \Gamma$ and $\Gamma_1 \subseteq \Gamma$, $\Gamma_0 \cup \Gamma_1 \subseteq \Gamma$. Hence Γ is inconsistent. □

4.10 Derivability and the Propositional Connectives

Proposition 4.22.

1. Both $\varphi \wedge \psi \vdash \varphi$ and $\varphi \wedge \psi \vdash \psi$.
2. $\varphi, \psi \vdash \varphi \wedge \psi$.

Proof. 1. Both sequents $\varphi \wedge \psi \Rightarrow \varphi$ and $\varphi \wedge \psi \Rightarrow \psi$ are **derivable**:

$$\frac{\varphi \Rightarrow \varphi}{\varphi \wedge \psi \Rightarrow \varphi} \wedge\text{L} \quad \frac{\psi \Rightarrow \psi}{\varphi \wedge \psi \Rightarrow \psi} \wedge\text{L}$$

2. Here is a **derivation** of the sequent $\varphi, \psi \Rightarrow \varphi \wedge \psi$:

$$\frac{\varphi \Rightarrow \varphi \quad \psi \Rightarrow \psi}{\varphi, \psi \Rightarrow \varphi \wedge \psi} \wedge\text{R}$$

□

Proposition 4.23.

1. $\varphi \vee \psi, \neg\varphi, \neg\psi$ is inconsistent.
2. Both $\varphi \vdash \varphi \vee \psi$ and $\psi \vdash \varphi \vee \psi$.

Proof. 1. We give a **derivation** of the sequent $\varphi \vee \psi, \neg\varphi, \neg\psi \Rightarrow$:

$$\frac{\frac{\frac{\varphi \Rightarrow \varphi}{\neg\varphi, \varphi \Rightarrow} \neg\text{L}}{\varphi, \neg\varphi, \neg\psi \Rightarrow} \quad \frac{\frac{\psi \Rightarrow \psi}{\neg\psi, \psi \Rightarrow} \neg\text{L}}{\psi, \neg\varphi, \neg\psi \Rightarrow}}{\varphi \vee \psi, \neg\varphi, \neg\psi \Rightarrow} \vee\text{L}$$

fol:seq:ppr:
sec
fol:seq:ppr:
prop:provability-land
fol:seq:ppr:
prop:provability-land-left
fol:seq:ppr:
prop:provability-land-right

fol:seq:ppr:
prop:provability-lor

(Recall that double inference lines indicate several weakening, contraction, and exchange inferences.)

2. Both sequents $\varphi \Rightarrow \varphi \vee \psi$ and $\psi \Rightarrow \varphi \vee \psi$ have **derivations**:

$$\frac{\varphi \Rightarrow \varphi}{\varphi \Rightarrow \varphi \vee \psi} \vee\text{R} \qquad \frac{\psi \Rightarrow \psi}{\psi \Rightarrow \varphi \vee \psi} \vee\text{R}$$

□

Proposition 4.24.

fol:seq:ppr:
prop:provability-lif
fol:seq:ppr:
prop:provability-lif-left
fol:seq:ppr:
prop:provability-lif-right

1. $\varphi, \varphi \rightarrow \psi \vdash \psi$.
2. Both $\neg\varphi \vdash \varphi \rightarrow \psi$ and $\psi \vdash \varphi \rightarrow \psi$.

Proof. 1. The sequent $\varphi \rightarrow \psi, \varphi \Rightarrow \psi$ is **derivable**:

$$\frac{\varphi \Rightarrow \varphi \quad \psi \Rightarrow \psi}{\varphi \rightarrow \psi, \varphi \Rightarrow \psi} \rightarrow\text{L}$$

2. Both sequents $\neg\varphi \Rightarrow \varphi \rightarrow \psi$ and $\psi \Rightarrow \varphi \rightarrow \psi$ are **derivable**:

$$\frac{\frac{\frac{\varphi \Rightarrow \varphi}{\neg\varphi, \varphi \Rightarrow} \neg\text{L}}{\varphi, \neg\varphi \Rightarrow} \text{XL}}{\varphi, \neg\varphi \Rightarrow \psi} \text{WR} \rightarrow\text{R} \qquad \frac{\frac{\psi \Rightarrow \psi}{\varphi, \psi \Rightarrow \psi} \text{WL}}{\psi \Rightarrow \varphi \rightarrow \psi} \rightarrow\text{R}$$

□

4.11 Derivability and the Quantifiers

fol:seq:qpr:
sec

fol:seq:qpr:
thm:strong-generalization

Theorem 4.25. *If c is a constant not occurring in Γ or $\varphi(x)$ and $\Gamma \vdash \varphi(c)$, then $\Gamma \vdash \forall x \varphi(x)$.*

Proof. Let π_0 be an **LK-derivation** of $\Gamma_0 \Rightarrow \varphi(c)$ for some finite $\Gamma_0 \subseteq \Gamma$. By adding a $\forall\text{R}$ inference, we obtain a proof of $\Gamma_0 \Rightarrow \forall x \varphi(x)$, since c does not occur in Γ or $\varphi(x)$ and thus the eigenvariable condition is satisfied. □

Proposition 4.26.

fol:seq:qpr:
prop:provability-quantifiers

1. $\varphi(t) \vdash \exists x \varphi(x)$.
2. $\forall x \varphi(x) \vdash \varphi(t)$.

Proof. 1. The sequent $\varphi(t) \Rightarrow \exists x \varphi(x)$ is **derivable**:

$$\frac{\varphi(t) \Rightarrow \varphi(t)}{\varphi(t) \Rightarrow \exists x \varphi(x)} \exists R$$

2. The sequent $\forall x \varphi(x) \Rightarrow \varphi(t)$ is **derivable**:

$$\frac{\varphi(t) \Rightarrow \varphi(t)}{\forall x \varphi(x) \Rightarrow \varphi(t)} \forall L$$

□

4.12 Soundness

explanation

A **derivation** system, such as the sequent calculus, is *sound* if it cannot **derive** things that do not actually hold. Soundness is thus a kind of guaranteed safety property for **derivation** systems. Depending on which proof theoretic property is in question, we would like to know for instance, that

fol:seq:sou:
sec

1. every **derivable sentence** is valid;
2. if a **sentence** is **derivable** from some others, it is also a consequence of them;
3. if a set of **sentences** is inconsistent, it is unsatisfiable.

These are important properties of a **derivation** system. If any of them do not hold, the **derivation** system is deficient—it would **derive** too much. Consequently, establishing the soundness of a **derivation** system is of the utmost importance.

Because all these proof-theoretic properties are defined via **derivability** in the sequent calculus of certain sequents, proving (1)–(3) above requires proving something about the semantic properties of **derivable** sequents. We will first define what it means for a sequent to be *valid*, and then show that every **derivable** sequent is valid. (1)–(3) then follow as corollaries from this result.

Definition 4.27. A **structure** \mathfrak{M} *satisfies* a sequent $\Gamma \Rightarrow \Delta$ iff either $\mathfrak{M} \not\models \varphi$ for some $\varphi \in \Gamma$ or $\mathfrak{M} \models \varphi$ for some $\varphi \in \Delta$.

A sequent is *valid* iff every **structure** \mathfrak{M} satisfies it.

Theorem 4.28 (Soundness). *If **LK** derives $\Theta \Rightarrow \Xi$, then $\Theta \Rightarrow \Xi$ is valid.*

fol:seq:sou:
thm:sequent-soundness

Proof. Let π be a **derivation** of $\Theta \Rightarrow \Xi$. We proceed by induction on the number of inferences n in π .

If the number of inferences is 0, then π consists only of an initial sequent. Every initial sequent $\varphi \Rightarrow \varphi$ is obviously valid, since for every \mathfrak{M} , either $\mathfrak{M} \not\models \varphi$ or $\mathfrak{M} \models \varphi$.

If the number of inferences is greater than 0, we distinguish cases according to the type of the lowermost inference. By induction hypothesis, we can assume

that the premises of that inference are valid, since the number of inferences in the proof of any premise is smaller than n .

First, we consider the possible inferences with only one premise.

1. The last inference is a weakening. Then $\Theta \Rightarrow \Xi$ is either $A, \Gamma \Rightarrow \Delta$ (if the last inference is WL) or $\Gamma \Rightarrow \Delta, \varphi$ (if it's WR), and the **derivation** ends in one of

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta \end{array}}{\varphi, \Gamma \Rightarrow \Delta} \text{WL} \qquad \frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta \end{array}}{\Gamma \Rightarrow \Delta, \varphi} \text{WR}$$

By induction hypothesis, $\Gamma \Rightarrow \Delta$ is valid, i.e., for every **structure** \mathfrak{M} , either there is some $\chi \in \Gamma$ such that $\mathfrak{M} \not\models \chi$ or there is some $\chi \in \Delta$ such that $\mathfrak{M} \models \chi$.

If $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, then $\chi \in \Theta$ as well since $\Theta = \varphi, \Gamma$, and so $\mathfrak{M} \not\models \chi$ for some $\chi \in \Theta$. Similarly, if $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$, as $\chi \in \Xi$, $\mathfrak{M} \models \chi$ for some $\chi \in \Xi$. Consequently, $\Theta \Rightarrow \Xi$ is valid.

2. The last inference is \neg L: Then the premise of the last inference is $\Gamma \Rightarrow \Delta, \varphi$ and the conclusion is $\neg\varphi, \Gamma \Rightarrow \Delta$, i.e., the **derivation** ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi \end{array}}{\neg\varphi, \Gamma \Rightarrow \Delta} \neg\text{L}$$

and $\Theta = \neg\varphi, \Gamma$ while $\Xi = \Delta$.

The induction hypothesis tells us that $\Gamma \Rightarrow \Delta, \varphi$ is valid, i.e., for every \mathfrak{M} , either (a) for some $\chi \in \Gamma$, $\mathfrak{M} \not\models \chi$, or (b) for some $\chi \in \Delta$, $\mathfrak{M} \models \chi$, or (c) $\mathfrak{M} \models \varphi$. We want to show that $\Theta \Rightarrow \Xi$ is also valid. Let \mathfrak{M} be a **structure**. If (a) holds, then there is $\chi \in \Gamma$ so that $\mathfrak{M} \not\models \varphi$, but $\varphi \in \Theta$ as well. If (b) holds, there is $\chi \in \Delta$ such that $\mathfrak{M} \models \chi$, but $\chi \in \Xi$ as well. Finally, if $\mathfrak{M} \models \varphi$, then $\mathfrak{M} \not\models \neg\varphi$. Since $\neg\varphi \in \Theta$, there is $\chi \in \Theta$ such that $\mathfrak{M} \not\models \chi$. Consequently, $\Theta \Rightarrow \Xi$ is valid.

3. The last inference is \neg R: Exercise.
4. The last inference is \wedge L: There are two variants: $\varphi \wedge \psi$ may be inferred on the left from φ or from ψ on the left side of the premise. In the first case, the π ends in

$$\frac{\begin{array}{c} \vdots \\ \varphi, \Gamma \Rightarrow \Delta \end{array}}{\varphi \wedge \psi, \Gamma \Rightarrow \Delta} \wedge L$$

and $\Theta = \varphi \wedge \psi, \Gamma$ while $\Xi = \Delta$. Consider a **structure** \mathfrak{M} . Since by induction hypothesis, $\varphi, \Gamma \Rightarrow \Delta$ is valid, (a) $\mathfrak{M} \not\models \varphi$, (b) $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, or (c) $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$. In case (a), $\mathfrak{M} \not\models \varphi \wedge \psi$, so there is $\chi \in \Theta$ (namely, $\varphi \wedge \psi$) such that $\mathfrak{M} \not\models \chi$. In case (b), there is $\chi \in \Gamma$ such that $\mathfrak{M} \not\models \alpha$, and $\chi \in \Theta$ as well. In case (c), there is $\chi \in \Delta$ such that $\mathfrak{M} \models \chi$, and $\chi \in \Xi$ as well since $\Xi = \Delta$. So in each case, \mathfrak{M} satisfies $\varphi \wedge \psi, \Gamma \Rightarrow \Delta$. Since \mathfrak{M} was arbitrary, $\Gamma \Rightarrow \Delta$ is valid. The case where $\varphi \wedge \psi$ is inferred from ψ is handled the same, changing φ to ψ .

5. The last inference is $\vee R$: There are two variants: $\varphi \vee \psi$ may be inferred on the right from φ or from ψ on the right side of the premise. In the first case, π ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi \end{array}}{\Gamma \Rightarrow \Delta, \varphi \vee \psi} \vee R$$

Now $\Theta = \Gamma$ and $\Xi = \Delta, \varphi \vee \psi$. Consider a **structure** \mathfrak{M} . Since $\Gamma \Rightarrow \Delta, \varphi$ is valid, (a) $\mathfrak{M} \models \varphi$, (b) $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, or (c) $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$. In case (a), $\mathfrak{M} \models \varphi \vee \psi$. In case (b), there is $\chi \in \Gamma$ such that $\mathfrak{M} \not\models \chi$. In case (c), there is $\chi \in \Delta$ such that $\mathfrak{M} \models \chi$. So in each case, \mathfrak{M} satisfies $\Gamma \Rightarrow \Delta, \varphi \vee \psi$, i.e., $\Theta \Rightarrow \Xi$. Since \mathfrak{M} was arbitrary, $\Theta \Rightarrow \Xi$ is valid. The case where $\varphi \vee \psi$ is inferred from ψ is handled the same, changing φ to ψ .

6. The last inference is $\rightarrow R$: Then π ends in

$$\frac{\begin{array}{c} \vdots \\ \varphi, \Gamma \Rightarrow \Delta, \varphi \end{array}}{\Gamma \Rightarrow \Delta, \varphi \rightarrow \psi} \rightarrow R$$

Again, the induction hypothesis says that the premise is valid; we want to show that the conclusion is valid as well. Let \mathfrak{M} be arbitrary. Since $\varphi, \Gamma \Rightarrow \Delta, \psi$ is valid, at least one of the following cases obtains: (a) $\mathfrak{M} \not\models \varphi$, (b) $\mathfrak{M} \models \psi$, (c) $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, or (d) $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$. In cases (a) and (b), $\mathfrak{M} \models \varphi \rightarrow \psi$ and so there is a $\chi \in \Delta, \varphi \rightarrow \psi$ such that $\mathfrak{M} \models \chi$. In case (c), for some $\chi \in \Gamma$, $\mathfrak{M} \not\models \chi$. In case (d), for some $\chi \in \Delta$, $\mathfrak{M} \models \chi$. In each case, \mathfrak{M} satisfies $\Gamma \Rightarrow \Delta, \varphi \rightarrow \psi$. Since \mathfrak{M} was arbitrary, $\Gamma \Rightarrow \Delta, \varphi \rightarrow \psi$ is valid.

7. The last inference is $\forall\text{L}$: Then there is a **formula** $\varphi(x)$ and a closed term t such that π ends in

$$\frac{\begin{array}{c} \vdots \\ \varphi(t), \Gamma \Rightarrow \Delta \end{array}}{\forall x \varphi(x), \Gamma \Rightarrow \Delta} \forall\text{L}$$

We want to show that the conclusion $\forall x \varphi(x), \Gamma \Rightarrow \Delta$ is valid. Consider a **structure** \mathfrak{M} . Since the premise $\varphi(t), \Gamma \Rightarrow \Delta$ is valid, (a) $\mathfrak{M} \not\models \varphi(t)$, (b) $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, or (c) $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$. In case (a), by [Proposition 1.53](#), if $\mathfrak{M} \models \forall x \varphi(x)$, then $\mathfrak{M} \models \varphi(t)$. Since $\mathfrak{M} \not\models \varphi(t)$, $\mathfrak{M} \not\models \forall x \varphi(x)$. In case (b) and (c), \mathfrak{M} also satisfies $\forall x \varphi(x), \Gamma \Rightarrow \Delta$. Since \mathfrak{M} was arbitrary, $\forall x \varphi(x), \Gamma \Rightarrow \Delta$ is valid.

8. The last inference is $\exists\text{R}$: Exercise.
9. The last inference is $\forall\text{R}$: Then there is a **formula** $\varphi(x)$ and a **constant symbol** a such that π ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi(a) \end{array}}{\Gamma \Rightarrow \Delta, \forall x \varphi(x)} \forall\text{R}$$

where the eigenvariable condition is satisfied, i.e., a does not occur in $\varphi(x), \Gamma$, or Δ . By induction hypothesis, the premise of the last inference is valid. We have to show that the conclusion is valid as well, i.e., that for any **structure** \mathfrak{M} , (a) $\mathfrak{M} \models \forall x \varphi(x)$, (b) $\mathfrak{M} \not\models \chi$ for some $\chi \in \Gamma$, or (c) $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$.

Suppose \mathfrak{M} is an arbitrary **structure**. If (b) or (c) holds, we are done, so suppose neither holds: for all $\chi \in \Gamma$, $\mathfrak{M} \models \chi$, and for all $\chi \in \Delta$, $\mathfrak{M} \not\models \chi$. We have to show that (a) holds, i.e., $\mathfrak{M} \models \forall x \varphi(x)$. By [Proposition 1.41](#), it suffices to show that $\mathfrak{M}, s \models \varphi(x)$ for all variable assignments s . So let s be an arbitrary variable assignment. Consider the structure \mathfrak{M}' which is just like \mathfrak{M} except $a^{\mathfrak{M}'} = s(x)$. By [Corollary 1.43](#), for any $\chi \in \Gamma$, $\mathfrak{M}' \models \chi$ since a does not occur in Γ , and for any $\chi \in \Delta$, $\mathfrak{M}' \not\models \chi$. But the premise is valid, so $\mathfrak{M}' \models \varphi(a)$. By [Proposition 1.40](#), $\mathfrak{M}', s \models \varphi(a)$, since $\varphi(a)$ is a sentence. Now $s \sim_x s$ with $s(x) = \text{Val}_s^{\mathfrak{M}'}(a)$, since we've defined \mathfrak{M}' in just this way. So [Proposition 1.45](#) applies, and we get $\mathfrak{M}', s \models \varphi(x)$. Since a does not occur in $\varphi(x)$, by [Proposition 1.42](#), $\mathfrak{M}, s \models \varphi(x)$. Since s was arbitrary, we've completed the proof that $\mathfrak{M}, s \models \varphi(x)$ for all variable assignments.

10. The last inference is $\exists\text{L}$: Exercise.

Now let's consider the possible inferences with two premises.

1. The last inference is a cut: then π ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi \end{array} \quad \begin{array}{c} \vdots \\ \varphi, \Pi \Rightarrow \Lambda \end{array}}{\Gamma, \Pi \Rightarrow \Delta, \Lambda} \text{Cut}$$

Let \mathfrak{M} be a **structure**. By induction hypothesis, the premises are valid, so \mathfrak{M} satisfies both premises. We distinguish two cases: (a) $\mathfrak{M} \not\models \varphi$ and (b) $\mathfrak{M} \models \varphi$. In case (a), in order for \mathfrak{M} to satisfy the left premise, it must satisfy $\Gamma \Rightarrow \Delta$. But then it also satisfies the conclusion. In case (b), in order for \mathfrak{M} to satisfy the right premise, it must satisfy $\Pi \Rightarrow \Lambda$. Again, \mathfrak{M} satisfies the conclusion.

2. The last inference is $\wedge R$. Then π ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi \end{array} \quad \begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \psi \end{array}}{\Gamma \Rightarrow \Delta, \varphi \wedge \psi} \wedge R$$

Consider a **structure** \mathfrak{M} . If \mathfrak{M} satisfies $\Gamma \Rightarrow \Delta$, we are done. So suppose it doesn't. Since $\Gamma \Rightarrow \Delta, \varphi$ is valid by induction hypothesis, $\mathfrak{M} \models \varphi$. Similarly, since $\Gamma \Rightarrow \Delta, \psi$ is valid, $\mathfrak{M} \models \psi$. But then $\mathfrak{M} \models \varphi \wedge \psi$.

3. The last inference is $\vee L$: Exercise.
4. The last inference is $\rightarrow L$. Then π ends in

$$\frac{\begin{array}{c} \vdots \\ \Gamma \Rightarrow \Delta, \varphi \end{array} \quad \begin{array}{c} \vdots \\ \psi, \Pi \Rightarrow \Lambda \end{array}}{\varphi \rightarrow \psi, \Gamma, \Pi \Rightarrow \Delta, \Lambda} \rightarrow L$$

Again, consider a **structure** \mathfrak{M} and suppose \mathfrak{M} doesn't satisfy $\Gamma, \Pi \Rightarrow \Delta, \Lambda$. We have to show that $\mathfrak{M} \not\models \varphi \rightarrow \psi$. If \mathfrak{M} doesn't satisfy $\Gamma, \Pi \Rightarrow \Delta, \Lambda$, it satisfies neither $\Gamma \Rightarrow \Delta$ nor $\Pi \Rightarrow \Lambda$. Since, $\Gamma \Rightarrow \Delta, \varphi$ is valid, we have $\mathfrak{M} \models \varphi$. Since $\psi, \Pi \Rightarrow \Lambda$ is valid, we have $\mathfrak{M} \not\models \psi$. But then $\mathfrak{M} \not\models \varphi \rightarrow \psi$, which is what we wanted to show.

□

Problem 4.5. Complete the proof of [Theorem 4.28](#).

fol:seq:sou:
cor:weak-soundness

Corollary 4.29. *If $\Gamma \vdash \varphi$ then φ is valid.*

fol:seq:sou:
cor:entailment-soundness

Corollary 4.30. *If $\Gamma \vdash \varphi$ then $\Gamma \vDash \varphi$.*

Proof. If $\Gamma \vdash \varphi$ then for some finite subset $\Gamma_0 \subseteq \Gamma$, there is a **derivation** of $\Gamma_0 \Rightarrow \varphi$. By **Theorem 4.28**, every **structure** \mathfrak{M} either makes some $\psi \in \Gamma_0$ false or makes φ true. Hence, if $\mathfrak{M} \models \Gamma$ then also $\mathfrak{M} \models \varphi$. \square

fol:seq:sou:
cor:consistency-soundness

Corollary 4.31. *If Γ is satisfiable, then it is consistent.*

Proof. We prove the contrapositive. Suppose that Γ is not consistent. Then there is a finite $\Gamma_0 \subseteq \Gamma$ and a **derivation** of $\Gamma_0 \Rightarrow \perp$. By **Theorem 4.28**, $\Gamma_0 \Rightarrow \perp$ is valid. In other words, for every **structure** \mathfrak{M} , there is $\chi \in \Gamma_0$ so that $\mathfrak{M} \not\models \chi$, and since $\Gamma_0 \subseteq \Gamma$, that χ is also in Γ . Thus, no \mathfrak{M} satisfies Γ , and Γ is not satisfiable. \square

4.13 Derivations with Identity predicate

fol:seq:ide:
sec

Derivations with **identity predicate** require additional initial sequents and inference rules.

Definition 4.32 (Initial sequents for $=$). If t is a closed term, then $\Rightarrow t = t$ is an initial sequent.

The rules for $=$ are (t_1 and t_2 are closed terms):

$$\boxed{\frac{t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_1)}{t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_2)} = \qquad \frac{t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_2)}{t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_1)} =}$$

Example 4.33. If s and t are closed terms, then $s = t, \varphi(s) \vdash \varphi(t)$:

$$\frac{\frac{\varphi(s) \Rightarrow \varphi(s)}{s = t, \varphi(s) \Rightarrow \varphi(s)} \text{WL}}{s = t, \varphi(s) \Rightarrow \varphi(t)} =$$

This may be familiar as the principle of substitutability of identicals, or Leibniz' Law.

LK proves that $=$ is symmetric and transitive:

$$\frac{\frac{\Rightarrow t_1 = t_1} {t_1 = t_2 \Rightarrow t_1 = t_1} \text{WL}}{t_1 = t_2 \Rightarrow t_2 = t_1} = \qquad \frac{\frac{t_1 = t_2 \Rightarrow t_1 = t_2} {t_2 = t_3, t_1 = t_2 \Rightarrow t_1 = t_2} \text{WL}}{t_2 = t_3, t_1 = t_2 \Rightarrow t_1 = t_3} = \text{XL}}{t_1 = t_2, t_2 = t_3 \Rightarrow t_1 = t_3}$$

In the proof on the left, the **formula** $x = t_1$ is our $\varphi(x)$. On the right, we take $\varphi(x)$ to be $t_1 = x$.

Problem 4.6. Give **derivations** of the following sequents:

1. $\Rightarrow \forall x \forall y ((x = y \wedge \varphi(x)) \rightarrow \varphi(y))$
2. $\exists x \varphi(x) \wedge \forall y \forall z ((\varphi(y) \wedge \varphi(z)) \rightarrow y = z) \Rightarrow \exists x (\varphi(x) \wedge \forall y (\varphi(y) \rightarrow y = x))$

4.14 Soundness with Identity predicate

fol:seq:sid:
sec

Proposition 4.34. **LK** with initial sequents and rules for identity is sound.

Proof. Initial sequents of the form $\Rightarrow t = t$ are valid, since for every **structure** \mathfrak{M} , $\mathfrak{M} \models t = t$. (Note that we assume the term t to be closed, i.e., it contains no variables, so variable assignments are irrelevant).

Suppose the last inference in a **derivation** is $=$. Then the premise is $t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_1)$ and the conclusion is $t_1 = t_2, \Gamma \Rightarrow \Delta, \varphi(t_2)$. Consider a **structure** \mathfrak{M} . We need to show that the conclusion is valid, i.e., if $\mathfrak{M} \models t_1 = t_2$ and $\mathfrak{M} \models \Gamma$, then either $\mathfrak{M} \models \chi$ for some $\chi \in \Delta$ or $\mathfrak{M} \models \varphi(t_2)$.

By induction hypothesis, the premise is valid. This means that if $\mathfrak{M} \models t_1 = t_2$ and $\mathfrak{M} \models \Gamma$ either (a) for some $\chi \in \Delta$, $\mathfrak{M} \models \chi$ or (b) $\mathfrak{M} \models \varphi(t_1)$. In case (a) we are done. Consider case (b). Let s be a variable assignment with $s(x) = \text{Val}^{\mathfrak{M}}(t_1)$. By [Proposition 1.40](#), $\mathfrak{M}, s \models \varphi(t_1)$. Since $s \sim_x s$, by [Proposition 1.45](#), $\mathfrak{M}, s \models \varphi(x)$. since $\mathfrak{M} \models t_1 = t_2$, we have $\text{Val}^{\mathfrak{M}}(t_1) = \text{Val}^{\mathfrak{M}}(t_2)$, and hence $s(x) = \text{Val}^{\mathfrak{M}}(t_2)$. By applying [Proposition 1.45](#) again, we also have $\mathfrak{M}, s \models \varphi(t_2)$. By [Proposition 1.40](#), $\mathfrak{M} \models \varphi(t_2)$. \square

Chapter 5

Natural Deduction

5.1 Rules and Derivations

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sec

Natural deduction systems are meant to closely parallel the informal reasoning used in mathematical proof (hence it is somewhat “natural”). Natural deduction proofs begin with assumptions. Inference rules are then applied. Assumptions are “discharged” by the \neg Intro, \rightarrow Intro, \forall Elim and \exists Elim inference rules, and the label of the discharged assumption is placed beside the inference for clarity.

explanation

Definition 5.1 (Initial Formula). An *initial formula* or *assumption* is any formula in the topmost position of any branch.

Derivations in natural deduction are certain trees of sentences, where the topmost sentences are assumptions, and if a sentence stands below one, two, or three other sequents, it must follow correctly by a rule of inference. The sentences at the top of the inference are called the *premises* and the sentence below the *conclusion* of the inference. The rules come in pairs, an introduction and an elimination rule for each logical operator. They introduce a logical operator in the conclusion or remove a logical operator from a premise of the rule. Some of the rules allow an assumption of a certain type to be discharged. To indicate which assumption is discharged by which inference, we also assign labels to both the assumption and the inference. This is indicated by writing the assumption as “[φ]^{*n*}”.

It is customary to consider rules for all logical operators, even for those (if any) that we consider as defined.

5.2 Propositional Rules

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sec

Rules for \wedge

$$\frac{\varphi \quad \psi}{\varphi \wedge \psi} \wedge\text{Intro} \qquad \frac{\varphi \wedge \psi}{\varphi} \wedge\text{Elim}$$

$$\frac{\varphi \wedge \psi}{\psi} \wedge\text{Elim}$$

Rules for \vee

$$\frac{\varphi}{\varphi \vee \psi} \vee\text{Intro}$$

$$\frac{\psi}{\varphi \vee \psi} \vee\text{Intro}$$

$$n \frac{\varphi \vee \psi \quad \begin{array}{c} [\varphi]^n \\ \vdots \\ \chi \end{array} \quad \begin{array}{c} [\psi]^n \\ \vdots \\ \chi \end{array}}{\chi} \vee\text{Elim}$$

Rules for \rightarrow

$$n \frac{\begin{array}{c} [\varphi]^n \\ \vdots \\ \psi \end{array}}{\varphi \rightarrow \psi} \rightarrow\text{Intro}$$

$$\frac{\varphi \rightarrow \psi \quad \varphi}{\psi} \rightarrow\text{Elim}$$

Rules for \neg

$$n \frac{\begin{array}{c} [\varphi]^n \\ \vdots \\ \perp \end{array}}{\neg\varphi} \neg\text{Intro}$$

$$\frac{\neg\varphi \quad \varphi}{\perp} \neg\text{Elim}$$

Rules for \perp

$$\frac{\perp}{\varphi} \perp_I \qquad \begin{array}{c} [\neg\varphi]^n \\ \vdots \\ \vdots \\ n \frac{\perp}{\varphi} \perp_C \end{array}$$

Note that \neg -Intro and \perp_C are very similar: The difference is that \neg -Intro derives a negated **sentence** $\neg\varphi$ but \perp_C a positive **sentence** φ .

5.3 Quantifier Rules

fol:ntd:qrl:
sec

Rules for \forall

$$\frac{\varphi(a)}{\forall x \varphi(x)} \forall\text{Intro} \qquad \frac{\forall x \varphi(x)}{\varphi(t)} \forall\text{Elim}$$

In the rules for \forall , t is a ground term (a term that does not contain any variables), and a is a **constant symbol** which does not occur in the conclusion $\forall x \varphi(x)$, or in any assumption which is **undischarged** in the **derivation** ending with the premise $\varphi(a)$. We call a the *eigenvariable* of the \forall Intro inference.

Rules for \exists

$$\frac{\varphi(t)}{\exists x \varphi(x)} \exists\text{Intro} \qquad \begin{array}{c} [\varphi(a)]^n \\ \vdots \\ \vdots \\ \vdots \\ n \frac{\exists x \varphi(x)}{\chi} \exists\text{Elim} \end{array}$$

Again, t is a ground term, and a is a constant which does not occur in the premise $\exists x \varphi(x)$, in the conclusion χ , or any assumption which is **undischarged** in the **derivations** ending with the two premises (other than the assumptions $\varphi(a)$). We call a the *eigenvariable* of the \exists Elim inference.

The condition that an eigenvariable neither occur in the premises nor in any assumption that is **undischarged** in the **derivations** leading to the premises for the \forall Intro or \exists Elim inference is called the *eigenvariable condition*.

explanation

We use the term “eigenvariable” even though a in the above rules is a constant. This has historical reasons.

In \exists Intro and \forall Elim there are no restrictions, and the term t can be anything, so we do not have to worry about any conditions. On the other hand, in the \exists Elim and \forall Intro rules, the eigenvariable condition requires that the constant symbol a does not occur anywhere in the conclusion or in an undischarged assumption. The condition is necessary to ensure that the system is sound, i.e., only derives sentences from undischarged assumptions from which the follow. Without this condition, the following would be allowed:

$$\frac{\exists x \varphi(x) \quad \frac{[\varphi(a)]^1}{\forall x \varphi(x)} * \forall \text{Intro}}{\forall x \varphi(x)} \exists \text{Elim}$$

However, $\exists x \varphi(x) \not\equiv \forall x \varphi(x)$.

5.4 Derivations

explanation

We’ve said what an assumption is, and we’ve given the rules of inference. **Derivations** in the sequent calculus are inductively generated from these: each **derivation** either is an assumption on its own, or consists of one, two, or three **derivations** followed by a correct inference.

fol:ntd:der:
sec

Definition 5.2 (Derivation). A *derivation* of a sentence φ from assumptions Γ is a tree of **sentences** satisfying the following conditions:

1. The topmost **sentences** of the tree are either in Γ or are **discharged** by an inference in the tree.
2. The bottommost **sentence** of the tree is φ .
3. Every **sentence** in the tree except φ is a premise of a correct application of an inference rule whose conclusion stands directly below that **sentence** in the tree.

We then say that φ is the *conclusion* of the **derivation** and that φ is *derivable* from Γ .

Example 5.3. Every assumption on its own is a **derivation**. So, e.g., χ by itself is a **derivation**, and so is θ by itself. We can obtain a new **derivation** from these by applying, say, the \wedge Intro rule,

$$\frac{\varphi \quad \psi}{\varphi \wedge \psi} \wedge \text{Intro}$$

These rules are meant to be general: we can replace the φ and ψ in it with any **sentences**, e.g., by χ and θ . Then the conclusion would be $\chi \wedge \theta$, and so

$$\frac{\chi \quad \theta}{\chi \wedge \theta} \wedge\text{Intro}$$

is a correct **derivation**. Of course, we can also switch the assumptions, so that θ plays the role of φ and χ that of ψ . Thus,

$$\frac{\theta \quad \chi}{\theta \wedge \chi} \wedge\text{Intro}$$

is also a correct derivation.

We can now apply another rule, say, \rightarrow Intro, which allows us to conclude a conditional and allows us to **discharge** any assumption that is identical to the conclusion of that conditional. So both of the following would be correct **derivations**:

$$1 \frac{\frac{[\chi]^1 \quad \theta}{\chi \wedge \theta} \wedge\text{Intro}}{\chi \rightarrow (\chi \wedge \theta)} \rightarrow\text{Intro} \quad 1 \frac{\frac{\chi \quad [\theta]^1}{\chi \wedge \theta} \wedge\text{Intro}}{\theta \rightarrow (\chi \wedge \theta)} \rightarrow\text{Intro}$$

5.5 Examples of Derivations

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sec

Example 5.4. Let's give a **derivation** of the **sentence** $(\varphi \wedge \psi) \rightarrow \varphi$.

We begin by writing the desired conclusion at the bottom of the **derivation**.

$$\overline{(\varphi \wedge \psi) \rightarrow \varphi}$$

Next, we need to figure out what kind of inference could result in a **sentence** of this form. The **main operator** of the conclusion is \rightarrow , so we'll try to arrive at the conclusion using the \rightarrow Intro rule. It is best to write down the assumptions involved and label the inference rules as you progress, so it is easy to see whether all assumptions have been **discharged** at the end of the proof.

$$1 \frac{\begin{array}{c} [\varphi \wedge \psi]^1 \\ \vdots \\ \varphi \end{array}}{(\varphi \wedge \psi) \rightarrow \varphi} \rightarrow\text{Intro}$$

We now need to fill in the steps from the assumption $\varphi \wedge \psi$ to φ . Since we only have one connective to deal with, \wedge , we must use the \wedge elim rule. This gives us the following proof:

$$1 \frac{\frac{[\varphi \wedge \psi]^1}{\varphi} \wedge\text{Elim}}{(\varphi \wedge \psi) \rightarrow \varphi} \rightarrow\text{Intro}$$

We now have a correct **derivation** of $(\varphi \wedge \psi) \rightarrow \varphi$.

Example 5.5. Now let's give a **derivation** of $(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)$.

We begin by writing the desired conclusion at the bottom of the derivation.

$$\overline{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)}$$

To find a logical rule that could give us this conclusion, we look at the logical connectives in the conclusion: \neg , \vee , and \rightarrow . We only care at the moment about the first occurrence of \rightarrow because it is the **main operator** of the **sentence** in the end-sequent, while \neg , \vee and the second occurrence of \rightarrow are inside the scope of another connective, so we will take care of those later. We therefore start with the \rightarrow Intro rule. A correct application must look as follows:

$$\begin{array}{c} [\neg\varphi \vee \psi]^1 \\ \vdots \\ \varphi \rightarrow \psi \\ 1 \frac{}{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)} \rightarrow\text{Intro} \end{array}$$

This leaves us with two possibilities to continue. Either we can keep working from the bottom up and look for another application of the \rightarrow Intro rule, or we can work from the top down and apply a \vee Elim rule. Let us apply the latter. We will use the assumption $\neg\varphi \vee \psi$ as the leftmost premise of \vee Elim. For a valid application of \vee Elim, the other two premises must be identical to the conclusion $\varphi \rightarrow \psi$, but each may be derived in turn from another assumption, namely the two disjuncts of $\neg\varphi \vee \psi$. So our **derivation** will look like this:

$$\begin{array}{c} [\neg\varphi]^2 \quad [\psi]^2 \\ \vdots \quad \vdots \\ [\neg\varphi \vee \psi]^1 \quad \varphi \rightarrow \psi \quad \varphi \rightarrow \psi \\ 2 \frac{}{\varphi \rightarrow \psi} \vee\text{Elim} \\ 1 \frac{}{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)} \rightarrow\text{Intro} \end{array}$$

In each of the two branches on the right, we want to **derive** $\varphi \rightarrow \psi$, which is best done using \rightarrow Intro.

$$\begin{array}{c} [\neg\varphi]^2, [\varphi]^3 \quad [\psi]^2, [\varphi]^4 \\ \vdots \quad \vdots \\ \psi \quad \psi \\ 3 \frac{}{\varphi \rightarrow \psi} \rightarrow\text{Intro} \quad 4 \frac{}{\varphi \rightarrow \psi} \rightarrow\text{Intro} \\ 2 \frac{[\neg\varphi \vee \psi]^1}{\varphi \rightarrow \psi} \vee\text{Elim} \\ 1 \frac{}{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)} \rightarrow\text{Intro} \end{array}$$

For the two missing parts of the **derivation**, we need **derivations** of ψ from $\neg\varphi$ and φ in the middle, and from φ and ψ on the left. Let's take the former

first. $\neg\varphi$ and φ are the two premises of \neg Elim:

$$\frac{[\neg\varphi]^2 \quad [\varphi]^3}{\perp} \neg\text{Elim}$$

$$\vdots$$

$$\psi$$

By using \perp_I , we can obtain ψ as a conclusion and complete the branch.

$$\frac{\frac{\frac{[\neg\varphi]^2 \quad [\varphi]^3}{\perp} \perp\text{Intro} \quad \frac{[\psi]^2, [\varphi]^4}{\vdots} \rightarrow\text{Intro}}{\frac{\perp}{\psi} \perp_I} \rightarrow\text{Intro} \quad \frac{[\psi]^2}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\frac{[\neg\varphi \vee \psi]^1}{\varphi \rightarrow \psi} \vee\text{Elim}} \rightarrow\text{Intro} \quad \frac{[\psi]^2, [\varphi]^4}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\frac{[\neg\varphi \vee \psi]^1}{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)} \rightarrow\text{Intro}} \rightarrow\text{Intro}$$

Let's now look at the rightmost branch. Here it's important to realize that the definition of **derivation** allows assumptions to be discharged but does not require them to be. In other words, if we can derive ψ from one of the assumptions φ and ψ without using the other, that's ok. And to **derive** ψ from ψ is trivial: ψ by itself is such a **derivation**, and no inferences are needed. So we can simply delete the assumption φ .

$$\frac{\frac{\frac{[\neg\varphi]^2 \quad [\varphi]^3}{\perp} \neg\text{Elim} \quad \frac{[\psi]^2}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\frac{\perp}{\psi} \perp_I} \rightarrow\text{Intro} \quad \frac{[\psi]^2}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\frac{[\neg\varphi \vee \psi]^1}{\varphi \rightarrow \psi} \vee\text{Elim}} \rightarrow\text{Intro} \quad \frac{[\psi]^2}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\frac{[\neg\varphi \vee \psi]^1}{(\neg\varphi \vee \psi) \rightarrow (\varphi \rightarrow \psi)} \rightarrow\text{Intro}} \rightarrow\text{Intro}$$

Note that in the finished **derivation**, the rightmost \rightarrow Intro inference does not actually discharge any assumptions.

Example 5.6. So far we have not needed the \perp_C rule. It is special in that it allows us to discharge an assumption that isn't a sub-formula of the conclusion of the rule. It is closely related to the \perp_I rule. In fact, the \perp_I rule is a special case of the \perp_C rule—there is a logic called “intuitionistic logic” in which only \perp_I is allowed. The \perp_C rule is a last resort when nothing else works. For instance, suppose we want to **derive** $\varphi \vee \neg\varphi$. Our usual strategy would be to attempt to **derive** $\varphi \vee \neg\varphi$ using \vee Intro. But this would require us to **derive** either φ or $\neg\varphi$ from no assumptions, and this can't be done. \perp_C to the rescue!

$$\frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \vdots \quad \perp}{1 \quad \frac{\perp}{\varphi \vee \neg\varphi} \perp_C} \perp_C$$

Now we're looking for a **derivation** of \perp from $\neg(\varphi \vee \neg\varphi)$. Since \perp is the conclusion of \neg -Elim we might try that:

$$\frac{\frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \vdots \quad \neg\varphi}{\neg\varphi} \quad \frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \vdots \quad \varphi}{\varphi}}{1 \quad \frac{\perp}{\varphi \vee \neg\varphi} \perp_C} \neg\text{-Elim}$$

Our strategy for finding a **derivation** of $\neg\varphi$ calls for an application of \neg -Intro:

$$\frac{\frac{[\neg(\varphi \vee \neg\varphi)]^1, [\varphi]^2 \quad \vdots \quad \perp}{2 \quad \frac{\perp}{\neg\varphi} \neg\text{-Intro}} \quad \frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \vdots \quad \varphi}{\varphi} \neg\text{-Elim}}{1 \quad \frac{\perp}{\varphi \vee \neg\varphi} \perp_C} \neg\text{-Elim}$$

Here, we can get \perp easily by applying \neg -Elim to the assumption $\neg(\varphi \vee \neg\varphi)$ and $\varphi \vee \neg\varphi$ which follows from our new assumption φ by \vee -Intro:

$$\frac{\frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \frac{[\varphi]^2}{\varphi \vee \neg\varphi} \vee\text{-Intro}}{\neg\text{-Elim}} \quad \frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \vdots \quad \varphi}{\varphi} \neg\text{-Elim}}{2 \quad \frac{\perp}{\neg\varphi} \neg\text{-Intro}} \neg\text{-Elim}}{1 \quad \frac{\perp}{\varphi \vee \neg\varphi} \perp_C} \perp_C$$

On the right side we use the same strategy, except we get φ by \perp_C :

$$\frac{\frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \frac{[\varphi]^2}{\varphi \vee \neg\varphi} \vee\text{-Intro}}{\neg\text{-Elim}} \quad \frac{[\neg(\varphi \vee \neg\varphi)]^1 \quad \frac{[\neg\varphi]^3}{\varphi \vee \neg\varphi} \vee\text{-Intro}}{\neg\text{-Elim}}}{2 \quad \frac{\perp}{\neg\varphi} \neg\text{-Intro} \quad 3 \quad \frac{\perp}{\varphi} \perp_C} \neg\text{-Elim}}{1 \quad \frac{\perp}{\varphi \vee \neg\varphi} \perp_C} \perp_C$$

Problem 5.1. Give **derivations** of the following:

1. $\neg(\varphi \rightarrow \psi) \rightarrow (\varphi \wedge \neg\psi)$
2. $(\varphi \rightarrow \chi) \vee (\psi \rightarrow \chi)$ from the assumption $(\varphi \wedge \psi) \rightarrow \chi$

5.6 Derivations with Quantifiers

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sec

Example 5.7. When dealing with quantifiers, we have to make sure not to violate the eigenvariable condition, and sometimes this requires us to play around with the order of carrying out certain inferences. In general, it helps to try and take care of rules subject to the eigenvariable condition first (they will be lower down in the finished proof).

Let's see how we'd give a **derivation** of the **formula** $\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)$. Starting as usual, we write

$$\overline{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)}$$

We start by writing down what it would take to justify that last step using the \rightarrow Intro rule.

$$\frac{\begin{array}{c} [\exists x \neg\varphi(x)]^1 \\ \vdots \\ \neg\forall x \varphi(x) \end{array}}{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)} \rightarrow\text{Intro}$$

Since there is no obvious rule to apply to $\neg\forall x \varphi(x)$, we will proceed by setting up the **derivation** so we can use the \exists Elim rule. Here we must pay attention to the eigenvariable condition, and choose a constant that does not appear in $\exists x \varphi(x)$ or any assumptions that it depends on. (Since no **constant symbols** appear, however, any choice will do fine.)

$$\frac{\begin{array}{c} [\neg\varphi(a)]^2 \\ \vdots \\ [\exists x \neg\varphi(x)]^1 \quad \neg\forall x \varphi(x) \end{array}}{\neg\forall x \varphi(x)} \exists\text{Elim} \quad \frac{\quad}{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)} \rightarrow\text{Intro}$$

In order to derive $\neg\forall x \varphi(x)$, we will attempt to use the \neg Intro rule: this requires that we derive a contradiction, possibly using $\forall x \varphi(x)$ as an additional assumption. Of coursem, this contradiction may involve the assumption $\neg\varphi(a)$ which will be discharged by the \rightarrow Intro inference. We can set it up as follows:

$$\frac{\begin{array}{c} [\neg\varphi(a)]^2, [\forall x \varphi(x)]^3 \\ \vdots \\ \perp \end{array}}{\neg\forall x \varphi(x)} \neg\text{Intro} \quad \frac{\quad}{\neg\forall x \varphi(x)} \exists\text{Elim} \quad \frac{\begin{array}{c} [\exists x \neg\varphi(x)]^1 \\ \vdots \\ \neg\forall x \varphi(x) \end{array}}{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)} \rightarrow\text{Intro}$$

It looks like we are close to getting a contradiction. The easiest rule to apply is the \forall Elim, which has no eigenvariable conditions. Since we can use any term we want to replace the universally quantified x , it makes the most sense to continue using a so we can reach a contradiction.

$$\frac{\frac{\frac{[\exists x \neg\varphi(x)]^1}{\neg\forall x \varphi(x)} \exists\text{Elim} \quad \frac{\frac{[\neg\varphi(a)]^2}{\perp} \neg\text{Intro} \quad \frac{[\forall x \varphi(x)]^3}{\varphi(a)} \forall\text{Elim}}{\neg\forall x \varphi(x)} \neg\text{Elim}}{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)} \rightarrow\text{Intro}}{\exists x \neg\varphi(x) \rightarrow \neg\forall x \varphi(x)} \rightarrow\text{Intro}$$

It is important, especially when dealing with quantifiers, to double check at this point that the eigenvariable condition has not been violated. Since the only rule we applied that is subject to the eigenvariable condition was \exists Elim, and the eigenvariable a does not occur in any assumptions it depends on, this is a correct derivation.

Example 5.8. Sometimes we may derive a **formula** from other **formulas**. In these cases, we may have undischarged assumptions. It is important to keep track of our assumptions as well as the end goal.

Let's see how we'd give a **derivation** of the **formula** $\exists x \chi(x, b)$ from the assumptions $\exists x (\varphi(x) \wedge \psi(x))$ and $\forall x (\psi(x) \rightarrow \chi(x, b))$. Starting as usual, we write the conclusion at the bottom.

$$\overline{\exists x \chi(x, b)}$$

We have two premises to work with. To use the first, i.e., try to find a **derivation** of $\exists x \chi(x, b)$ from $\exists x (\varphi(x) \wedge \psi(x))$ we would use the \exists Elim rule. Since it has an eigenvariable condition, we will apply that rule first. We get the following:

$$1 \frac{\frac{[\varphi(a) \wedge \psi(a)]^1}{\vdots} \quad \exists x \chi(x, b)}{\exists x (\varphi(x) \wedge \psi(x)) \quad \exists x \chi(x, b)} \exists\text{Elim}$$

The two assumptions we are working with share ψ . It may be useful at this point to apply \wedge Elim to separate out $\psi(a)$.

$$1 \frac{\frac{[\varphi(a) \wedge \psi(a)]^1}{\psi(a)} \wedge\text{Elim} \quad \exists x \chi(x, b)}{\exists x (\varphi(x) \wedge \psi(x)) \quad \exists x \chi(x, b)} \exists\text{Elim}$$

The second assumption we have to work with is $\forall x (\psi(x) \rightarrow \chi(x, b))$. Since there is no eigenvariable condition we can instantiate x with the **constant symbol** a using \forall Elim to get $\psi(a) \rightarrow \chi(a, b)$. We now have both $\psi(a) \rightarrow \chi(a, b)$ and $\psi(a)$. Our next move should be a straightforward application of the \rightarrow Elim rule.

$$\begin{array}{c}
 \frac{\frac{\forall x (\psi(x) \rightarrow \chi(x, b))}{\psi(a) \rightarrow \chi(a, b)} \forall\text{Elim} \quad \frac{[\varphi(a) \wedge \psi(a)]^1}{\psi(a)} \wedge\text{Elim}}{\chi(a, b)} \rightarrow\text{Elim} \\
 \vdots \\
 \frac{\exists x (\varphi(x) \wedge \psi(x)) \quad \exists x \chi(x, b)}{\exists x \chi(x, b)} \exists\text{Elim}
 \end{array}$$

We are so close! One application of \exists Intro and we have reached our goal.

$$\begin{array}{c}
 \frac{\frac{\forall x (\psi(x) \rightarrow \chi(x, b))}{\psi(a) \rightarrow \chi(a, b)} \forall\text{Elim} \quad \frac{[\varphi(a) \wedge \psi(a)]^1}{\psi(a)} \wedge\text{Elim}}{\chi(a, b)} \rightarrow\text{Elim} \\
 \frac{\exists x (\varphi(x) \wedge \psi(x)) \quad \frac{\chi(a, b)}{\exists x \chi(x, b)} \exists\text{Intro}}{\exists x \chi(x, b)} \exists\text{Elim}
 \end{array}$$

Since we ensured at each step that the eigenvariable conditions were not violated, we can be confident that this is a correct derivation.

Example 5.9. Give a **derivation** of the **formula** $\neg\forall x \varphi(x)$ from the assumptions $\forall x \varphi(x) \rightarrow \exists y \psi(y)$ and $\neg\exists y \psi(y)$. Starting as usual, we write the target **formula** at the bottom.

$$\overline{\neg\forall x \varphi(x)}$$

The last line of the **derivation** is a negation, so let's try using \neg Intro. This will require that we figure out how to **derive** a contradiction.

$$\begin{array}{c}
 [\forall x \varphi(x)]^1 \\
 \vdots \\
 \perp \\
 \frac{}{\neg\forall x \varphi(x)} \neg\text{Intro}
 \end{array}$$

So far so good. We can use \forall Elim but it's not obvious if that will help us get to our goal. Instead, let's use one of our assumptions. $\forall x \varphi(x) \rightarrow \exists y \psi(y)$

together with $\forall x \varphi(x)$ will allow us to use the \rightarrow Elim rule.

$$\frac{\forall x \varphi(x) \rightarrow \exists y \psi(y) \quad [\forall x \varphi(x)]^1}{\exists y \psi(y)} \rightarrow\text{Elim}$$

$$\vdots$$

$$1 \frac{\perp}{\neg \forall x \varphi(x)} \neg\text{Intro}$$

We now have one final assumption to work with, and it looks like this will help us reach a contradiction by using \neg Elim.

$$\frac{\neg \exists y \psi(y) \quad \frac{\forall x \varphi(x) \rightarrow \exists y \psi(y) \quad [\forall x \varphi(x)]^1}{\exists y \psi(y)} \rightarrow\text{Elim}}{1 \frac{\perp}{\neg \forall x \varphi(x)} \neg\text{Intro}} \neg\text{Elim}$$

Problem 5.2. Give **derivations** of the following:

1. $\exists y \varphi(y) \rightarrow \psi$ from the assumption $\forall x (\varphi(x) \rightarrow \psi)$
2. $\exists x (\varphi(x) \rightarrow \forall y \varphi(y))$

5.7 Proof-Theoretic Notions

explanation

Just as we've defined a number of important semantic notions (validity, entailment, satisfiability), we now define corresponding *proof-theoretic notions*. These are not defined by appeal to satisfaction of **sentences** in **structures**, but by appeal to the **derivability** or **non-derivability** of certain **sentences** from others. It was an important discovery, due to Gödel, that these notions coincide. That they do is the content of the *completeness theorem*.

fol:ntd:ptn:
sec

Definition 5.10 (Theorems). A **sentence** φ is a *theorem* if there is a **derivation** of φ in natural deduction in which all assumptions are **discharged**. We write $\vdash \varphi$ if φ is a theorem and $\not\vdash \varphi$ if it is not.

Definition 5.11 (Derivability). A **sentence** φ is *derivable* from a set of **sentences** Γ , $\Gamma \vdash \varphi$, if there is a **derivation** with conclusion φ and in which every assumption is either **discharged** or is in Γ . If φ is not **derivable** from Γ we write $\Gamma \not\vdash \varphi$.

Definition 5.12 (Consistency). A set of **sentences** Γ is *inconsistent* iff $\Gamma \vdash \perp$. If Γ is not inconsistent, i.e., if $\Gamma \not\vdash \perp$, we say it is *consistent*.

Proposition 5.13 (Reflexivity). If $\varphi \in \Gamma$, then $\Gamma \vdash \varphi$.

fol:ntd:ptn:
prop:reflexivity

Proof. The assumption φ by itself is a **derivation** of φ where every **undischarged** assumption (i.e., φ) is in Γ . \square

*fol.mtd:ptn:
prop:monotony*

Proposition 5.14 (Monotony). *If $\Gamma \subseteq \Delta$ and $\Gamma \vdash \varphi$, then $\Delta \vdash \varphi$.*

Proof. Any **derivation** of φ from Γ is also a **derivation** of φ from Δ . □

*fol.mtd:ptn:
prop:transitivity*

Proposition 5.15 (Transitivity). *If $\Gamma \vdash \varphi$ for every $\varphi \in \Delta$ and $\Delta \vdash \psi$, then $\Gamma \vdash \psi$.*

Proof. If $\Delta \vdash \psi$, then there is a **derivation** δ_0 of ψ with all **undischarged** assumptions in Δ . We show that $\Gamma \vdash \psi$ by induction on the number n of **undischarged** assumptions in δ_0 .

If $n = 0$, then δ_0 has no **undischarged** assumptions, and so also counts as a **derivation** of ψ from Γ .

Otherwise, pick an **undischarged** assumption φ in δ_0 and let Δ_1 be the remaining **undischarged** assumptions. We obtain the **derivation** δ_1 :

$$\frac{\begin{array}{c} \Delta_1, [\varphi]^1 \\ \vdots \\ \vdots \delta_0 \\ \vdots \\ \psi \end{array}}{\varphi \rightarrow \psi} \rightarrow \text{Intro}$$

Since the number of **undischarged** assumptions in δ_1 is $n - 1$, the inductive hypothesis applies: there is a **derivation** δ_2 of $\varphi \rightarrow \psi$ from Γ . Since $\Gamma \vdash \varphi$ there is also a **derivation** δ_3 of φ from Γ . Now consider:

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \vdots \delta_2 \\ \vdots \\ \varphi \rightarrow \psi \end{array} \quad \begin{array}{c} \Gamma \\ \vdots \\ \vdots \delta_3 \\ \vdots \\ \varphi \end{array}}{\psi} \rightarrow \text{Elim}$$

This shows $\Gamma \vdash \psi$. □

*fol.mtd:ptn:
prop:incons*

Proposition 5.16. *Γ is inconsistent iff $\Gamma \vdash \varphi$ for every **sentence** φ .*

Proof. Exercise. □

Problem 5.3. Prove Proposition 5.16

*fol.mtd:ptn:
prop:proves-compact*

Proposition 5.17 (Compactness).

1. *If $\Gamma \vdash \varphi$ then there is a finite subset $\Gamma_0 \subseteq \Gamma$ such that $\Gamma_0 \vdash \varphi$.*
2. *If every finite subset of Γ is consistent, then Γ is consistent.*

Proof. 1. If $\Gamma \vdash \varphi$, then there is a **derivation** δ of φ from Γ . Let Γ_0 be the set of **undischarged** assumptions of δ . Since any **derivation** is finite, Γ_0 can only contain finitely many **sentences**. So, δ is a **derivation** of φ from a finite $\Gamma_0 \subseteq \Gamma$.

2. This is the contrapositive of (1) for the special case $\varphi \equiv \perp$. □

5.8 Derivability and Consistency

We will now establish a number of properties of the **derivability** relation. They are independently interesting, but each will play a role in the proof of the completeness theorem.

Proposition 5.18. *If $\Gamma \vdash \varphi$ and $\Gamma \cup \{\varphi\}$ is inconsistent, then Γ is inconsistent.*

Proof. Let the **derivation** of φ from Γ be δ_1 and the **derivation** of \perp from $\Gamma \cup \{\varphi\}$ be δ_2 . We can then **derive**:

$$\frac{\begin{array}{c} \Gamma, [\varphi]^1 \\ \vdots \\ \delta_2 \\ \vdots \\ \perp \\ \hline \neg\varphi \quad \text{-Intro} \end{array}}{\perp} \quad \frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta_1 \\ \vdots \\ \varphi \\ \hline \varphi \quad \text{-Elim} \end{array}}{\perp} \quad \text{-Elim}$$

In the new **derivation**, the assumption φ is **discharged**, so it is a **derivation** from Γ . □

Proposition 5.19. *$\Gamma \vdash \varphi$ iff $\Gamma \cup \{\neg\varphi\}$ is inconsistent.*

Proof. First suppose $\Gamma \vdash \varphi$, i.e., there is a **derivation** δ_0 of φ from **undischarged** assumptions Γ . We obtain a **derivation** of \perp from $\Gamma \cup \{\neg\varphi\}$ as follows:

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta_0 \\ \vdots \\ \varphi \\ \hline \varphi \quad \text{-Elim} \end{array}}{\perp} \quad \neg\varphi$$

Now assume $\Gamma \cup \{\neg\varphi\}$ is inconsistent, and let δ_1 be the corresponding derivation of \perp from **undischarged** assumptions in $\Gamma \cup \{\neg\varphi\}$. We obtain a **derivation** of φ from Γ alone by using \perp_C :

$$\frac{\begin{array}{c} \Gamma, [\neg\varphi]^1 \\ \vdots \\ \delta_1 \\ \vdots \\ \perp \\ \hline \perp \quad \perp_C \end{array}}{\varphi}$$

□

Problem 5.4. Prove that $\Gamma \vdash \neg\varphi$ iff $\Gamma \cup \{\varphi\}$ is inconsistent.

Proposition 5.20. *If $\Gamma \vdash \varphi$ and $\neg\varphi \in \Gamma$, then Γ is inconsistent.*

*fol:ntd:prv:
prop:explicit-inc*

Proof. Suppose $\Gamma \vdash \varphi$ and $\neg\varphi \in \Gamma$. Then there is a **derivation** δ of φ from Γ . Consider this simple application of the \neg Elim rule:

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta \\ \vdots \\ \varphi \end{array}}{\perp} \neg\text{Elim}$$

Since $\neg\varphi \in \Gamma$, all **undischarged** assumptions are in Γ , this shows that $\Gamma \vdash \perp$. \square

fol.ntd.prv:
prop.provability-exhaustive

Proposition 5.21. *If $\Gamma \cup \{\varphi\}$ and $\Gamma \cup \{\neg\varphi\}$ are both inconsistent, then Γ is inconsistent.*

Proof. There are **derivations** δ_1 and δ_2 of \perp from $\Gamma \cup \{\varphi\}$ and \perp from $\Gamma \cup \{\neg\varphi\}$, respectively. We can then **derive**

$$\frac{\begin{array}{c} \Gamma, [\neg\varphi]^2 \\ \vdots \\ \delta_2 \\ \vdots \\ \perp \end{array} \quad \begin{array}{c} \Gamma, [\varphi]^1 \\ \vdots \\ \delta_1 \\ \vdots \\ \perp \end{array}}{\perp} \frac{\begin{array}{c} 2 \quad \neg\neg\varphi \\ \neg\text{Intro} \end{array} \quad \begin{array}{c} 1 \quad \neg\varphi \\ \neg\text{Intro} \end{array}}{\neg\text{Elim}}$$

Since the assumptions φ and $\neg\varphi$ are **discharged**, this is a **derivation** of \perp from Γ alone. Hence Γ is inconsistent. \square

5.9 Derivability and the Propositional Connectives

fol.ntd.ppr:
sec

Proposition 5.22.

fol.ntd.ppr:
prop.provability-land
fol.ntd.ppr:
prop.provability-land-left
fol.ntd.ppr:
prop.provability-land-right

1. Both $\varphi \wedge \psi \vdash \varphi$ and $\varphi \wedge \psi \vdash \psi$
2. $\varphi, \psi \vdash \varphi \wedge \psi$.

Proof. 1. We can **derive** both

$$\frac{\varphi \wedge \psi}{\varphi} \wedge\text{Elim} \quad \frac{\varphi \wedge \psi}{\psi} \wedge\text{Elim}$$

2. We can **derive**:

$$\frac{\varphi \quad \psi}{\varphi \wedge \psi} \wedge\text{Intro}$$

\square

Proposition 5.23.

*fol.ntd.ppr:
prop:provability-lor*

1. $\varphi \vee \psi, \neg\varphi, \neg\psi$ is inconsistent.
2. Both $\varphi \vdash \varphi \vee \psi$ and $\psi \vdash \varphi \vee \psi$.

Proof. 1. Consider the following derivation:

$$\frac{\frac{\frac{\varphi \vee \psi}{1} \quad \frac{\frac{\neg\varphi \quad [\varphi]^1}{\perp} \neg\text{Elim}}{\perp} \vee\text{Elim}}{\perp} \neg\text{Elim}}{\perp} \vee\text{Elim}}$$

This is a derivation of \perp from undischarged assumptions $\varphi \vee \psi, \neg\varphi$, and $\neg\psi$.

2. We can derive both

$$\frac{\varphi}{\varphi \vee \psi} \vee\text{Intro} \qquad \frac{\psi}{\varphi \vee \psi} \vee\text{Intro}$$

□

Proposition 5.24.

*fol.ntd.ppr:
prop:provability-lif*

1. $\varphi, \varphi \rightarrow \psi \vdash \psi$.
2. Both $\neg\varphi \vdash \varphi \rightarrow \psi$ and $\psi \vdash \varphi \rightarrow \psi$.

*fol.ntd.ppr:
prop:provability-lif-left*

*fol.ntd.ppr:
prop:provability-lif-right*

Proof. 1. We can derive:

$$\frac{\frac{\varphi \rightarrow \psi \quad \psi}{\psi} \rightarrow\text{Elim}}{\psi} \rightarrow\text{Elim}$$

2. This is shown by the following two derivations:

$$\frac{\frac{\frac{\frac{\neg\varphi \quad [\varphi]^1}{\perp} \neg\text{Elim}}{\perp} \perp_I}{\varphi \rightarrow \psi} \rightarrow\text{Intro}}{\varphi \rightarrow \psi} \rightarrow\text{Intro} \qquad \frac{\psi}{\varphi \rightarrow \psi} \rightarrow\text{Intro}$$

Note that \rightarrow Intro may, but does not have to, discharge the assumption φ .

□

5.10 Derivability and the Quantifiers

fol:ntd:qpr:
sec

fol:ntd:qpr:
thm:strong-generalization

Theorem 5.25. *If c is a constant not occurring in Γ or $\varphi(x)$ and $\Gamma \vdash \varphi(c)$, then $\Gamma \vdash \forall x \varphi(x)$.*

Proof. Let δ be a **derivation** of $\varphi(c)$ from Γ . By adding a \forall Intro inference, we obtain a proof of $\forall x \varphi(x)$. Since c does not occur in Γ or $\varphi(x)$, the eigenvariable condition is satisfied. \square

fol:ntd:qpr:
prop:provability-quantifiers

Proposition 5.26.

1. $\varphi(t) \vdash \exists x \varphi(x)$.

2. $\forall x \varphi(x) \vdash \varphi(t)$.

Proof. 1. The following is a **derivation** of $\exists x \varphi(x)$ from $\varphi(t)$:

$$\frac{\varphi(t)}{\exists x \varphi(x)} \exists\text{Intro}$$

2. The following is a **derivation** of $\varphi(t)$ from $\forall x \varphi(x)$:

$$\frac{\forall x \varphi(x)}{\varphi(t)} \forall\text{Elim}$$

\square

5.11 Soundness

fol:ntd:sou:
sec

A **derivation** system, such as natural deduction, is *sound* if it cannot **derive** explanation things that do not actually follow. Soundness is thus a kind of guaranteed safety property for **derivation** systems. Depending on which proof theoretic property is in question, we would like to know for instance, that

1. every **derivable sentence** is valid;
2. if a **sentence** is **derivable** from some others, it is also a consequence of them;
3. if a set of **sentences** is inconsistent, it is unsatisfiable.

These are important properties of a **derivation** system. If any of them do not hold, the **derivation** system is deficient—it would **derive** too much. Consequently, establishing the soundness of a **derivation** system is of the utmost importance.

fol:ntd:sou:
thm:soundness

Theorem 5.27 (Soundness). *If φ is **derivable** from the **undischarged assumptions** Γ , then $\Gamma \models \varphi$.*

Proof. Let δ be a **derivation** of φ . We proceed by induction on the number of inferences in δ .

For the induction basis we show the claim if the number of inferences is 0. In this case, δ consists only of an initial **formula**. Every initial **formula** φ is an **undischarged** assumption, and as such, any **structure** \mathfrak{M} that satisfies all of the **undischarged** assumptions of the proof also satisfies φ .

Now for the inductive step. Suppose that δ contains n inferences. The premise(s) of the lowermost inference are **derived** using sub-**derivations**, each of which contains fewer than n inferences. We assume the induction hypothesis: The premises of the last inference follow from the **undischarged** assumptions of the sub-**derivations** ending in those premises. We have to show that φ follows from the **undischarged** assumptions of the entire proof.

We distinguish cases according to the type of the lowermost inference. First, we consider the possible inferences with only one premise.

1. Suppose that the last inference is \neg -Intro: The **derivation** has the form

$$\begin{array}{c} \Gamma, [\varphi]^n \\ \vdots \\ \vdots \delta_1 \\ \vdots \\ \perp \\ \hline n \frac{}{\neg\varphi} \neg\text{Intro} \end{array}$$

By inductive hypothesis, \perp follows from the **undischarged** assumptions $\Gamma \cup \{\varphi\}$ of δ_1 . Consider a **structure** \mathfrak{M} . We need to show that, if $\mathfrak{M} \models \Gamma$, then $\mathfrak{M} \models \neg\varphi$. Suppose for reductio that $\mathfrak{M} \models \Gamma$, but $\mathfrak{M} \not\models \neg\varphi$, i.e., $\mathfrak{M} \models \varphi$. This would mean that $\mathfrak{M} \models \Gamma \cup \{\varphi\}$. This is contrary to our inductive hypothesis. So, $\mathfrak{M} \models \neg\varphi$.

2. The last inference is \wedge Elim: There are two variants: φ or ψ may be inferred from the premise $\varphi \wedge \psi$. Consider the first case. The derivation δ looks like this:

$$\begin{array}{c} \Gamma \\ \vdots \\ \vdots \delta_1 \\ \vdots \\ \varphi \wedge \psi \\ \hline \varphi \wedge\text{Elim} \end{array}$$

By inductive hypothesis, $\varphi \wedge \psi$ follows from the **undischarged** assumptions Γ of δ_1 . Consider a **structure** \mathfrak{M} . We need to show that, if $\mathfrak{M} \models \Gamma$, then $\mathfrak{M} \models \varphi$. Suppose $\mathfrak{M} \models \Gamma$. By our inductive hypothesis ($\Gamma \models \varphi \vee \psi$), we know that $\mathfrak{M} \models \varphi \wedge \psi$. By definition, $\mathfrak{M} \models \varphi \wedge \psi$ iff $\mathfrak{M} \models \varphi$ and $\mathfrak{M} \models \psi$. (The case where ψ is inferred from $\varphi \wedge \psi$ is handled similarly.)

3. The last inference is \vee Intro: There are two variants: $\varphi \vee \psi$ may be inferred from the premise φ or the premise ψ . Consider the first case. The derivation has the form

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta_1 \\ \vdots \\ \varphi \end{array}}{\varphi \vee \psi} \vee\text{Intro}$$

By inductive hypothesis, φ follows from the **undischarged** assumptions Γ of δ_1 . Consider a **structure** \mathfrak{M} . We need to show that, if $\mathfrak{M} \models \Gamma$, then $\mathfrak{M} \models \varphi \vee \psi$. Suppose $\mathfrak{M} \models \Gamma$; then $\mathfrak{M} \models \varphi$ since $\Gamma \models \varphi$ (the inductive hypothesis). So it must also be the case that $\mathfrak{M} \models \varphi \vee \psi$. (The case where $\varphi \vee \psi$ is inferred from ψ is handled similarly.)

4. The last inference is \rightarrow Intro: $\varphi \rightarrow \psi$ is inferred from a subproof with assumption φ and conclusion ψ , i.e.,

$${}^n \frac{\begin{array}{c} \Gamma, [\varphi]^n \\ \vdots \\ \delta_1 \\ \vdots \\ \psi \end{array}}{\varphi \rightarrow \psi} \rightarrow\text{Intro}$$

By inductive hypothesis, ψ follows from the **undischarged** assumptions of δ_1 , i.e., $\Gamma \cup \{\varphi\} \models \psi$. Consider a **structure** \mathfrak{M} . The **undischarged** assumptions of δ are just Γ , since φ is discharged at the last inference. So we need to show that $\Gamma \models \varphi \rightarrow \psi$. For reductio, suppose that for some **structure** \mathfrak{M} , $\mathfrak{M} \models \Gamma$ but $\mathfrak{M} \not\models \varphi \rightarrow \psi$. So, $\mathfrak{M} \models \varphi$ and $\mathfrak{M} \not\models \psi$. But by hypothesis, ψ is a consequence of $\Gamma \cup \{\varphi\}$, i.e., $\mathfrak{M} \models \psi$, which is a contradiction. So, $\Gamma \models \varphi \rightarrow \psi$.

5. The last inference is \perp_I : Here, δ ends in

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta_1 \\ \vdots \\ \perp \end{array}}{\varphi} \perp_I$$

By induction hypothesis, $\Gamma \models \perp$. We have to show that $\Gamma \models \varphi$. Suppose not; then for some \mathfrak{M} we have $\mathfrak{M} \models \Gamma$ and $\mathfrak{M} \not\models \varphi$. But we always have $\mathfrak{M} \models \perp$, so this would mean that $\Gamma \not\models \perp$, contrary to the induction hypothesis.

6. The last inference is \perp_C : Exercise.
7. The last inference is \forall Intro: Then δ has the form

$$\frac{\begin{array}{c} \Gamma \\ \vdots \\ \delta_1 \\ \vdots \\ \varphi(a) \end{array}}{\forall x \varphi(x)} \forall\text{Intro}$$

The premise $\varphi(a)$ is a consequence of the **undischarged** assumptions Γ by induction hypothesis. Consider some structure, \mathfrak{M} , such that $\mathfrak{M} \models \Gamma$. We need to show that $\mathfrak{M} \models \forall x \varphi(x)$. Since $\forall x \varphi(x)$ is a **sentence**, this means we have to show that for every variable assignment s , $\mathfrak{M}, s \models \varphi(x)$ (**Proposition 1.41**). Since Γ consists entirely of sentences, $\mathfrak{M}, s \models \psi$ for all $\psi \in \Gamma$ by **Definition 1.34**. Let \mathfrak{M}' be like \mathfrak{M} except that $a^{\mathfrak{M}'} = s(x)$. Since a does not occur in Γ , $\mathfrak{M}' \models \Gamma$ by **Corollary 1.43**. Since $\Gamma \vDash A(a)$, $\mathfrak{M}' \models A(a)$. Since $\varphi(a)$ is a **sentence**, $\mathfrak{M}, s \models \varphi(a)$ by **Proposition 1.40**. $\mathfrak{M}', s \models \varphi(x)$ iff $\mathfrak{M}' \models \varphi(a)$ by **Proposition 1.45** (recall that $\varphi(a)$ is just $\varphi(x)[a/x]$). So, $\mathfrak{M}', s \models \varphi(x)$. Since a does not occur in $\varphi(x)$, by **Proposition 1.42**, $\mathfrak{M}, s \models \varphi(x)$. But s was an arbitrary variable assignment, so $\mathfrak{M} \models \forall x \varphi(x)$.

8. The last inference is \exists Intro: Exercise.
9. The last inference is \forall Elim: Exercise.

Now let's consider the possible inferences with several premises: \forall Elim, \wedge Intro, \rightarrow Elim, and \exists Elim.

1. The last inference is \wedge Intro. $\varphi \wedge \psi$ is inferred from the premises φ and ψ and δ has the form

$$\frac{\begin{array}{c} \Gamma_1 \\ \vdots \\ \delta_1 \\ \vdots \\ \varphi \end{array} \quad \begin{array}{c} \Gamma_2 \\ \vdots \\ \delta_2 \\ \vdots \\ \psi \end{array}}{\varphi \wedge \psi} \wedge\text{Intro}$$

By induction hypothesis, φ follows from the **undischarged** assumptions Γ_1 of δ_1 and ψ follows from the **undischarged** assumptions Γ_2 of δ_2 . The **undischarged** assumptions of δ are $\Gamma_1 \cup \Gamma_2$, so we have to show that $\Gamma_1 \cup \Gamma_2 \vDash \varphi \wedge \psi$. Consider a **structure** \mathfrak{M} with $\mathfrak{M} \models \Gamma_1 \cup \Gamma_2$. Since $\mathfrak{M} \models \Gamma_1$, it must be the case that $\mathfrak{M} \models \varphi$ as $\Gamma_1 \vDash \varphi$, and since $\mathfrak{M} \models \Gamma_2$, $\mathfrak{M} \models \psi$ since $\Gamma_2 \vDash \psi$. Together, $\mathfrak{M} \models \varphi \wedge \psi$.

2. The last inference is \forall Elim: Exercise.

3. The last inference is \rightarrow Elim. ψ is inferred from the premises $\varphi \rightarrow \psi$ and φ . The derivation δ looks like this:

$$\frac{\begin{array}{c} \Gamma_1 \\ \vdots \\ \delta_1 \\ \vdots \\ \varphi \rightarrow \psi \end{array} \quad \begin{array}{c} \Gamma_2 \\ \vdots \\ \delta_2 \\ \vdots \\ \varphi \end{array}}{\psi} \rightarrow\text{Elim}$$

By induction hypothesis, $\varphi \rightarrow \psi$ follows from the **undischarged** assumptions Γ_1 of δ_1 and φ follows from the **undischarged** assumptions Γ_2 of δ_2 . Consider a **structure** \mathfrak{M} . We need to show that, if $\mathfrak{M} \models \Gamma_1 \cup \Gamma_2$, then $\mathfrak{M} \models \psi$. Suppose $\mathfrak{M} \models \Gamma_1 \cup \Gamma_2$. Since $\Gamma_1 \models \varphi \rightarrow \psi$, $\mathfrak{M} \models \varphi \rightarrow \psi$. Since $\Gamma_2 \models \varphi$, we have $\mathfrak{M} \models \varphi$. This means that $\mathfrak{M} \models \psi$ (For if $\mathfrak{M} \not\models \psi$, since $\mathfrak{M} \models \varphi$, we'd have $\mathfrak{M} \not\models \varphi \rightarrow \psi$, contradicting $\mathfrak{M} \models \varphi \rightarrow \psi$).

4. The last inference is \neg Elim: Exercise.
5. The last inference is \exists Elim: Exercise.

□

Problem 5.5. Complete the proof of [Theorem 5.27](#).

fol:ntd:sou:
cor:weak-soundness

Corollary 5.28. *If $\vdash \varphi$, then φ is valid.*

fol:ntd:sou:
cor:consistency-soundness

Corollary 5.29. *If Γ is satisfiable, then it is consistent.*

Proof. We prove the contrapositive. Suppose that Γ is not consistent. Then $\Gamma \vdash \perp$, i.e., there is a **derivation** of \perp from **undischarged** assumptions in Γ . By [Theorem 5.27](#), any **structure** \mathfrak{M} that satisfies Γ must satisfy \perp . Since $\mathfrak{M} \not\models \perp$ for every **structure** \mathfrak{M} , no \mathfrak{M} can satisfy Γ , i.e., Γ is not satisfiable. □

5.12 Derivations with Identity predicate

fol:ntd:ide:
sec

Derivations with identity predicate require additional inference rules.

$\frac{}{t = t} =\text{Intro}$	$\frac{t_1 = t_2 \quad \varphi(t_1)}{\varphi(t_2)} =\text{Elim}$ $\frac{t_1 = t_2 \quad \varphi(t_2)}{\varphi(t_1)} =\text{Elim}$
--------------------------------	---

In the above rules, t , t_1 , and t_2 are closed terms. The =Intro rule allows us to **derive** any identity statement of the form $t = t$ outright, from no assumptions.

Example 5.30. If s and t are closed terms, then $\varphi(s), s = t \vdash \varphi(t)$:

$$\frac{s = t \quad \varphi(s)}{\varphi(t)} =\text{Elim}$$

This may be familiar as the “principle of substitutability of identicals,” or Leibniz’ Law.

Problem 5.6. Prove that $=$ is both symmetric and transitive, i.e., give **derivations** of $\forall x \forall y (x = y \rightarrow y = x)$ and $\forall x \forall y \forall z ((x = y \wedge y = z) \rightarrow x = z)$

Example 5.31. We **derive** the **sentence**

$$\forall x \forall y ((\varphi(x) \wedge \varphi(y)) \rightarrow x = y)$$

from the **sentence**

$$\exists x \forall y (\varphi(y) \rightarrow y = x)$$

We develop the **derivation** backwards:

$$\begin{array}{c} \exists x \forall y (\varphi(y) \rightarrow y = x) \quad [\varphi(a) \wedge \varphi(b)]^1 \\ \vdots \\ a = b \\ \frac{1 \quad \frac{\frac{(\varphi(a) \wedge \varphi(b)) \rightarrow a = b}{\forall y ((\varphi(a) \wedge \varphi(y)) \rightarrow a = y)} \rightarrow\text{Intro}}{\forall y ((\varphi(a) \wedge \varphi(y)) \rightarrow a = y)} \forall\text{Intro}}{\forall x \forall y ((\varphi(x) \wedge \varphi(y)) \rightarrow x = y)} \forall\text{Intro} \end{array}$$

We’ll now have to use the main assumption: since it is an existential **formula**, we use $\exists\text{Elim}$ to **derive** the intermediary conclusion $a = b$.

$$\begin{array}{c} [\forall y (\varphi(y) \rightarrow y = c)]^2 \\ [\varphi(a) \wedge \varphi(b)]^1 \\ \vdots \\ \frac{2 \quad \frac{\frac{\frac{\frac{\frac{\exists x \forall y (\varphi(y) \rightarrow y = x)}{a = b} \exists\text{Elim}}{(\varphi(a) \wedge \varphi(b)) \rightarrow a = b} \rightarrow\text{Intro}}{\forall y ((\varphi(a) \wedge \varphi(y)) \rightarrow a = y)} \forall\text{Intro}}{\forall x \forall y ((\varphi(x) \wedge \varphi(y)) \rightarrow x = y)} \forall\text{Intro}}{a = b} \end{array}$$

The sub-**derivation** on the top right is completed by using its assumptions to show that $a = c$ and $b = c$. This requires two separate **derivations**. The derivation for $a = c$ is as follows:

$$\frac{\frac{[\forall y (\varphi(y) \rightarrow y = c)]^2}{\varphi(a) \rightarrow a = c} \forall\text{Elim} \quad \frac{[\varphi(a) \wedge \varphi(b)]^1}{\varphi(a)} \wedge\text{Elim}}{a = c} \rightarrow\text{Elim}$$

From $a = c$ and $b = c$ we derive $a = b$ by $=\text{Elim}$.

Problem 5.7. Give derivations of the following formulas:

1. $\forall x \forall y ((x = y \wedge \varphi(x)) \rightarrow \varphi(y))$
2. $\exists x \varphi(x) \wedge \forall y \forall z ((\varphi(y) \wedge \varphi(z)) \rightarrow y = z) \rightarrow \exists x (\varphi(x) \wedge \forall y (\varphi(y) \rightarrow y = x))$

5.13 Soundness with Identity predicate

fol:ntd:sid:
sec

Proposition 5.32. *Natural deduction with rules for $=$ is sound.*

Proof. Any formula of the form $t = t$ is valid, since for every structure \mathfrak{M} , $\mathfrak{M} \models t = t$. (Note that we assume the term t to be ground, i.e., it contains no variables, so variable assignments are irrelevant).

Suppose the last inference in a derivation is $=\text{Elim}$, i.e., the derivation has the following form:

$$\frac{\begin{array}{c} \Gamma_1 \\ \vdots \\ \delta_1 \\ \vdots \\ t_1 = t_2 \end{array} \quad \begin{array}{c} \Gamma_2 \\ \vdots \\ \delta_2 \\ \vdots \\ \varphi(t_1) \end{array}}{\varphi(t_2)} =\text{Elim}$$

The premises $t_1 = t_2$ and $\varphi(t_1)$ are derived from undischarged assumptions Γ_1 and Γ_2 , respectively. We want to show that $\varphi(t_2)$ follows from $\Gamma_1 \cup \Gamma_2$. Consider a structure \mathfrak{M} with $\mathfrak{M} \models \Gamma_1 \cup \Gamma_2$. By induction hypothesis, $\mathfrak{M} \models \varphi(t_1)$ and $\mathfrak{M} \models t_1 = t_2$. Therefore, $\text{Val}^{\mathfrak{M}}(t_1) = \text{Val}^{\mathfrak{M}}(t_2)$. Let s be any variable assignment, and s' be the x -variant given by $s'(x) = \text{Val}^{\mathfrak{M}}(t_1) = \text{Val}^{\mathfrak{M}}(t_2)$. By Proposition 1.45, $\mathfrak{M}, s \models \varphi(t_1)$ iff $\mathfrak{M}, s' \models \varphi(x)$ iff $\mathfrak{M}, s \models \varphi(t_2)$. Since $\mathfrak{M} \models \varphi(t_1)$, we have $\mathfrak{M} \models \varphi(t_2)$. \square

Chapter 6

Axiomatic Derivations

6.1 Rules and Derivations

fol:axd:rul:
sec Axiomatic **derivations** are perhaps the simplest proof system for logic. A explanation **derivation** is just a sequence of **formulas**. To count as a **derivation**, every **formula** in the sequence must either be an instance of an axiom, or must follow from one or more **formulas** that precede it in the sequence by a rule of inference. A **derivation** **derives** its last **formula**.

Definition 6.1 (*Derivability*). If Γ is a set of **formulas** of \mathcal{L} then a **derivation** from Γ is a finite sequence $\varphi_1, \dots, \varphi_n$ of **formulas** where for each $i \leq n$ one of the following holds:

1. $\varphi_i \in \Gamma$; or
2. φ_i is an axiom; or
3. φ_i follows from some φ_j (and φ_k) with $j < i$ (and $k < i$) by a rule of inference.

What counts as a correct **derivation** depends on which inference rules we allow (and of course what we take to be axioms). And an inference rule is an if-then statement that tells us that, under certain conditions, a step A_i in is a correct inference step.

Definition 6.2 (*Rule of inference*). A *rule of inference* gives a sufficient condition for what counts as a correct inference step in a **derivation** from Γ .

For instance, since any one-element sequence φ with $\varphi \in \Gamma$ trivially counts as a **derivation**, the following might be a very simple rule of inference:

If $\varphi \in \Gamma$, then φ is always a correct inference step in any **derivation** from Γ .

Similarly, if φ is one of the axioms, then φ by itself is a **derivation**, and so this is also a rule of inference:

If φ is an axiom, then φ is a correct inference step.

It gets more interesting if the rule of inference appeals to **formulas** that appear before the step considered. The following rule is called *modus ponens*:

If $\psi \rightarrow \varphi$ and ψ occur higher up in the **derivation**, then φ is a correct inference step.

If this is the only rule of inference, then our definition of **derivation** above amounts to this: $\varphi_1, \dots, \varphi_n$ is a **derivation** iff for each $i \leq n$ one of the following holds:

1. $\varphi_i \in \Gamma$; or
2. φ_i is an axiom; or
3. for some $j < i$, φ_j is $\psi \rightarrow \varphi_i$, and for some $k < i$, φ_k is ψ .

The last clause says that φ_i follows from φ_j (ψ) and φ_k ($\psi \rightarrow \varphi_i$) by modus ponens. If we can go from 1 to n , and each time we find a **formula** φ_i that is either in Γ , an axiom, or which a rule of inference tells us that it is a correct inference step, then the entire sequence counts as a correct **derivation**.

Definition 6.3 (Derivability). A **formula** φ is *derivable* from Γ , written $\Gamma \vdash \varphi$, if there is a **derivation** from Γ ending in φ .

Definition 6.4 (Theorems). A **formula** φ is a *theorem* if there is a **derivation** of φ from the empty set. We write $\vdash \varphi$ if φ is a theorem and $\not\vdash \varphi$ if it is not.

6.2 Axiom and Rules for the Propositional Connectives

fol:axd:prp:
sec

Definition 6.5 (Axioms). The set of Ax_0 of *axioms* for the propositional

connectives comprises all **formulas** of the following forms:

$$\text{fol:axd:prp:} \quad (\varphi \wedge \psi) \rightarrow \varphi \quad (6.1)$$

$$\text{fol:axd:prp:} \quad (\varphi \wedge \psi) \rightarrow \psi \quad (6.2)$$

$$\text{fol:axd:prp:} \quad \varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi)) \quad (6.3)$$

$$\text{fol:axd:prp:} \quad \varphi \rightarrow (\varphi \vee \psi) \quad (6.4)$$

$$\text{fol:axd:prp:} \quad \varphi \rightarrow (\psi \vee \varphi) \quad (6.5)$$

$$\text{fol:axd:prp:} \quad (\varphi \rightarrow \chi) \rightarrow ((\psi \rightarrow \chi) \rightarrow ((\varphi \vee \psi) \rightarrow \chi)) \quad (6.6)$$

$$\text{fol:axd:prp:} \quad \varphi \rightarrow (\psi \rightarrow \varphi) \quad (6.7)$$

$$\text{fol:axd:prp:} \quad (\varphi \rightarrow (\psi \rightarrow \chi)) \rightarrow ((\varphi \rightarrow \psi) \rightarrow (\varphi \rightarrow \chi)) \quad (6.8)$$

$$\text{fol:axd:prp:} \quad (\varphi \rightarrow \psi) \rightarrow ((\varphi \rightarrow \neg\psi) \rightarrow \neg\varphi) \quad (6.9)$$

$$\text{fol:axd:prp:} \quad \neg\varphi \rightarrow (\varphi \rightarrow \psi) \quad (6.10)$$

$$\text{fol:axd:prp:} \quad \top \quad (6.11)$$

$$\text{fol:axd:prp:} \quad \perp \rightarrow \varphi \quad (6.12)$$

$$\text{fol:axd:prp:} \quad (\varphi \rightarrow \perp) \rightarrow \neg\varphi \quad (6.13)$$

$$\text{fol:axd:prp:} \quad \neg\neg\varphi \rightarrow \varphi \quad (6.14)$$

ax:dne

Definition 6.6 (Modus ponens). If ψ and $\psi \rightarrow \varphi$ already occur in a derivation, then φ is a correct inference step.

We'll abbreviate the rule modus ponens as “MP.”

6.3 Axioms and Rules for Quantifiers

fol:axd:qua:
sec

Definition 6.7 (Axioms for quantifiers). The *axioms* governing quantifiers are all instances of the following:

$$\text{fol:axd:qua:} \quad \forall x \psi \rightarrow \psi(t), \quad (6.15)$$

$$\text{fol:axd:qua:} \quad \psi(t) \rightarrow \exists x \psi. \quad (6.16)$$

ax:q2

for any ground term t .

Definition 6.8 (Rules for quantifiers).

If $\psi \rightarrow \varphi(a)$ already occurs in the **derivation** and a does not occur in Γ or ψ , then $\psi \rightarrow \forall x \varphi(x)$ is a correct inference step.

If $\varphi(a) \rightarrow \psi$ already occurs in the **derivation** and a does not occur in Γ or ψ , then $\exists x \varphi(x) \rightarrow \psi$ is a correct inference step.

We'll abbreviate either of these by “QR.”

6.4 Examples of Derivations

fol:axd:pro:
sec

Example 6.9. Suppose we want to prove $(\neg\theta \vee \alpha) \rightarrow (\theta \rightarrow \alpha)$. Clearly, this is not an instance of any of our axioms, so we have to use the MP rule to **derive** it. Our only rule is MP, which given φ and $\varphi \rightarrow \psi$ allows us to justify ψ . One strategy would be to use eq. (6.6) with φ being $\neg\theta$, ψ being α , and χ being $\theta \rightarrow \alpha$, i.e., the instance

$$(\neg\theta \rightarrow (\theta \rightarrow \alpha)) \rightarrow ((\alpha \rightarrow (\theta \rightarrow \alpha)) \rightarrow ((\neg\theta \vee \alpha) \rightarrow (\theta \rightarrow \alpha))).$$

Why? Two applications of MP yield the last part, which is what we want. And we easily see that $\neg\theta \rightarrow (\theta \rightarrow \alpha)$ is an instance of eq. (6.10), and $\alpha \rightarrow (\theta \rightarrow \alpha)$ is an instance of eq. (6.7). So our derivation is:

1. $\neg\theta \rightarrow (\theta \rightarrow \alpha)$ eq. (6.7)
2. $(\neg\theta \rightarrow (\theta \rightarrow \alpha)) \rightarrow$
 $((\alpha \rightarrow (\theta \rightarrow \alpha)) \rightarrow ((\neg\theta \vee \alpha) \rightarrow (\theta \rightarrow \alpha)))$ eq. (6.6)
3. $((\alpha \rightarrow (\theta \rightarrow \alpha)) \rightarrow ((\neg\theta \vee \alpha) \rightarrow (\theta \rightarrow \alpha)))$ 1, 2, MP
4. $\alpha \rightarrow (\theta \rightarrow \alpha)$ eq. (6.7)
5. $(\neg\theta \vee \alpha) \rightarrow (\theta \rightarrow \alpha)$ 3, 4, MP

Example 6.10. Let's try to find a **derivation** of $\theta \rightarrow \theta$. It is not an instance of an axiom, so we have to use MP to **derive** it. eq. (6.7) is an axiom of the form $\varphi \rightarrow \psi$ to which we could apply MP. To be useful, of course, the ψ which MP would justify as a correct step in this case would have to be $\theta \rightarrow \theta$, since this is what we want to **derive**. That means φ would also have to be θ , i.e., we might look at this instance of eq. (6.7):

fol:axd:pro:
ex:identity

$$\theta \rightarrow (\theta \rightarrow \theta)$$

In order to apply MP, we would also need to justify the corresponding second premise, namely φ . But in our case, that would be θ , and we won't be able to **derive** θ by itself. So we need a different strategy.

The other axiom involving just \rightarrow is eq. (6.8), i.e.,

$$(\varphi \rightarrow (\psi \rightarrow \chi)) \rightarrow ((\varphi \rightarrow \psi) \rightarrow (\varphi \rightarrow \chi))$$

We could get to the last nested conditional by applying MP twice. Again, that would mean that we want an instance of eq. (6.8) where $\varphi \rightarrow \chi$ is $\theta \rightarrow \theta$, the **formula** we are aiming for. Then of course, φ and χ are both θ . How should we pick ψ so that both $\varphi \rightarrow (\psi \rightarrow \chi)$ and $\varphi \rightarrow \psi$, i.e., in our case $\theta \rightarrow (\psi \rightarrow \theta)$ and $\theta \rightarrow \psi$, are also **derivable**? Well, the first of these is already an instance of eq. (6.7), whatever we decide ψ to be. And $\theta \rightarrow \psi$ would be another instance of eq. (6.7) if ψ were $(\theta \rightarrow \theta)$. So, our derivation is:

1. $\theta \rightarrow ((\theta \rightarrow \theta) \rightarrow \theta)$ eq. (6.7)
2. $(\theta \rightarrow ((\theta \rightarrow \theta) \rightarrow \theta)) \rightarrow$
 $((\theta \rightarrow (\theta \rightarrow \theta)) \rightarrow (\theta \rightarrow \theta))$ eq. (6.8)
3. $(\theta \rightarrow (\theta \rightarrow \theta)) \rightarrow (\theta \rightarrow \theta)$ 1, 2, MP
4. $\theta \rightarrow (\theta \rightarrow \theta)$ eq. (6.7)
5. $\theta \rightarrow \theta$ 3, 4, MP

Example 6.11. Sometimes we want to show that there is a derivation of some formula from some other formulas Γ . For instance, let's show that we can derive $\varphi \rightarrow \chi$ from $\Gamma = \{\varphi \rightarrow \psi, \psi \rightarrow \chi\}$.

1. $\varphi \rightarrow \psi$ HYP
2. $\psi \rightarrow \chi$ HYP
3. $(\psi \rightarrow \chi) \rightarrow (\varphi \rightarrow (\psi \rightarrow \chi))$ eq. (6.7)
4. $\varphi \rightarrow (\psi \rightarrow \chi)$ 2, 3, MP
5. $(\varphi \rightarrow (\psi \rightarrow \chi)) \rightarrow$
 $((\varphi \rightarrow \psi) \rightarrow (\varphi \rightarrow \chi))$ eq. (6.8)
6. $((\varphi \rightarrow \psi) \rightarrow (\varphi \rightarrow \chi))$ 4, 5, MP
7. $\varphi \rightarrow \chi$ 1, 6, MP

The lines labelled “HYP” (for “hypothesis”) indicate that the formula on that line is an element of Γ .

Proposition 6.12. *If $\Gamma \vdash \varphi \rightarrow \psi$ and $\Gamma \vdash \psi \rightarrow \chi$, then $\Gamma \vdash \varphi \rightarrow \chi$*

Proof. Suppose $\Gamma \vdash \varphi \rightarrow \psi$ and $\Gamma \vdash \psi \rightarrow \chi$. Then there is a derivation of $\varphi \rightarrow \psi$ from Γ ; and a derivation of $\psi \rightarrow \chi$ from Γ as well. Combine these into a single derivation by concatenating them. Now add lines 3–7 of the derivation in the preceding example. This is a derivation of $\varphi \rightarrow \chi$ —which is the last line of the new derivation—from Γ . Note that the justifications of lines 4 and 7 remain valid if the reference to line number 2 is replaced by reference to the last line of the derivation of $\varphi \rightarrow \psi$, and reference to line number 1 by reference to the last line of the derivation of $\psi \rightarrow \chi$. \square

Problem 6.1. Show that the following hold by exhibiting derivations from the axioms:

1. $(\varphi \wedge \psi) \rightarrow (\psi \wedge \varphi)$
2. $((\varphi \wedge \psi) \rightarrow \chi) \rightarrow (\varphi \rightarrow (\psi \rightarrow \chi))$
3. $\neg(\varphi \vee \psi) \rightarrow \neg\varphi$

6.5 Derivations with Quantifiers

fol:axd:prq:
sec

Example 6.13. Let us give a derivation of $(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \forall x (\varphi(x) \wedge \psi(x))$.

First, note that

$$(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \forall x \varphi(x)$$

is an instance of eq. (6.1), and

$$\forall x \varphi(x) \rightarrow \varphi(a)$$

of eq. (6.15). So, by Proposition 6.12, we know that

$$(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \varphi(a)$$

is *derivable*. Likewise, since

$$\begin{aligned} (\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \forall y \psi(y) \quad \text{and} \\ \forall y \psi(y) \rightarrow \psi(a) \end{aligned}$$

are instances of eq. (6.2) and eq. (6.15), respectively,

$$(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \psi(a)$$

is derivable by Proposition 6.12. Using an appropriate instance of eq. (6.3) and two applications of MP, we see that

$$(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow (\varphi(a) \wedge \psi(a))$$

is derivable. We can now apply QR to obtain

$$(\forall x \varphi(x) \wedge \forall y \psi(y)) \rightarrow \forall x (\varphi(x) \wedge \psi(x)).$$

6.6 Proof-Theoretic Notions

explanation

Just as we've defined a number of important semantic notions (validity, entailment, satisfiability), we now define corresponding *proof-theoretic notions*. These are not defined by appeal to satisfaction of **sentences** in **structures**, but by appeal to the **derivability** or **non-derivability** of certain formulas. It was an important discovery, due to Gödel, that these notions coincide. That they do is the content of the *completeness theorem*.

fol:axd:ptn:
sec

The proof-theoretic notions for propositional logic are similar.

Definition 6.14 (Derivability). A formula φ is *derivable* from Γ , written $\Gamma \vdash \varphi$, if there is a **derivation** from Γ ending in φ .

Definition 6.15 (Theorems). A formula φ is a *theorem* if there is a **derivation** of φ from the empty set. We write $\vdash \varphi$ if φ is a theorem and $\not\vdash \varphi$ if it is not.

Definition 6.16 (Consistency). A set Γ of formulas is *consistent* if and only if $\Gamma \not\vdash \perp$; it is *inconsistent* otherwise.

fol:axd:ptn: **Proposition 6.17** (Reflexivity). *prop:reflexivity* If $\varphi \in \Gamma$, then $\Gamma \vdash \varphi$.

Proof. The formula φ by itself is a derivation of φ from Γ . □

fol:axd:ptn: **Proposition 6.18** (Monotony). *prop:monotony* If $\Gamma \subseteq \Delta$ and $\Gamma \vdash \varphi$, then $\Delta \vdash \varphi$.

Proof. Any derivation of φ from Γ is also a derivation of φ from Δ . □

fol:axd:ptn: **Proposition 6.19** (Transitivity). *prop:transitivity* If $\Gamma \vdash \varphi$ for every $\varphi \in \Delta$ and $\Delta \vdash \psi$, then $\Gamma \vdash \psi$.

Proof. Suppose $\Delta \vdash \psi$. Then there is a derivation $\psi_1, \dots, \psi_l = \psi$ from Δ . Some of the steps in that derivation will be correct because of a rule which refers to a prior line ψ_i that includes a formula $\psi_i \in \Delta$. Let $\varphi_1, \dots, \varphi_n$ be all the formulas in Δ so referenced. For each one of them, by hypothesis, there is a derivation of φ_i from Γ , i.e., a derivation $\varphi_i^1, \dots, \varphi_i^{k_i} = \varphi_i$ where every φ_i^j is an axiom, an element of Γ , or correct by a rule of inference. Now consider the sequence

$$\varphi_1^1, \dots, \varphi_1^{k_1}, \dots, \varphi_n^1, \dots, \varphi_n^{k_n}, \psi_1, \dots, \psi_l = \psi.$$

This is a correct derivation of ψ from Γ since each $\psi_i \in \Delta$ needed to justify the inferences in the part ψ_1, \dots, ψ_l is now justified itself. □

fol:axd:ptn: **Proposition 6.20.** *prop:incons* Γ is inconsistent iff $\Gamma \vdash \varphi$ for every sentence φ .

Proof. Exercise. □

Problem 6.2. Prove Proposition 6.20

fol:axd:ptn: **Proposition 6.21** (Compactness). *prop:proves-compact*

1. If $\Gamma \vdash \varphi$ then there is a finite subset $\Gamma_0 \subseteq \Gamma$ such that $\Gamma_0 \vdash \varphi$.
2. If every finite subset of Γ is consistent, then Γ is consistent.

Proof. 1. If $\Gamma \vdash \varphi$, then there is a finite sequence of formulas $\varphi_1, \dots, \varphi_n$ so that $\varphi \equiv \varphi_n$ and each φ_i is either a logical axiom, an element of Γ or follows from previous formulas by modus ponens. Take Γ_0 to be those φ_i which are in Γ . Then the derivation is likewise a derivation from Γ_0 , and so $\Gamma_0 \vdash \varphi$.

2. This is the contrapositive of (1) for the special case $\varphi \equiv \perp$. □

6.7 The Deduction Theorem

As we've seen, giving **derivations** in an axiomatic system is cumbersome, and **derivations** may be hard to find. Rather than actually write out long lists of **formulas**, it is generally easier to argue that such **derivations** exist, by making use of a few simple results. We've already established three such results: [Proposition 6.17](#) says we can always assert that $\Gamma \vdash \varphi$ when we know that $\varphi \in \Gamma$. [Proposition 6.18](#) says that if $\Gamma \vdash \varphi$ then also $\Gamma \cup \{\psi\} \vdash \varphi$. And [Proposition 6.19](#) implies that if $\Gamma \vdash \varphi$ and $\varphi \vdash \psi$, then $\Gamma \vdash \psi$. Here's another simple result, a “meta”-version of modus ponens:

fol:axd:ded:
sec

Proposition 6.22. *If $\Gamma \vdash \varphi$ and $\Gamma \vdash \varphi \rightarrow \psi$, then $\Gamma \vdash \psi$.*

fol:axd:ded:
prop:mp

Proof. We have that $\{\varphi, \varphi \rightarrow \psi\} \vdash \psi$:

1. φ Hyp.
2. $\varphi \rightarrow \psi$ Hyp.
3. ψ 1, 2, MP

By [Proposition 6.19](#), $\Gamma \vdash \psi$. □

The most important result we'll use in this context is the deduction theorem:

Theorem 6.23 (Deduction Theorem). *$\Gamma \cup \{\varphi\} \vdash \psi$ if and only if $\Gamma \vdash \varphi \rightarrow \psi$.*

fol:axd:ded:
thm:deduction-thm

Proof. The “if” direction is immediate. If $\Gamma \vdash \varphi \rightarrow \psi$ then also $\Gamma \cup \{\varphi\} \vdash \varphi \rightarrow \psi$ by [Proposition 6.18](#). Also, $\Gamma \cup \{\varphi\} \vdash \varphi$ by [Proposition 6.17](#). So, by [Proposition 6.22](#), $\Gamma \cup \{\varphi\} \vdash \psi$.

For the “only if” direction, we proceed by induction on the length of the **derivation** of ψ from $\Gamma \cup \{\varphi\}$.

For the induction basis, we prove the claim for every **derivation** of length 1. A **derivation** of ψ from $\Gamma \cup \{\varphi\}$ of length 1 consists of ψ by itself; and if it is correct ψ is either $\in \Gamma \cup \{\varphi\}$ or is an axiom. If $\psi \in \Gamma$ or is an axiom, then $\Gamma \vdash \psi$. We also have that $\Gamma \vdash \psi \rightarrow (\varphi \rightarrow \psi)$ by [eq. \(6.7\)](#), and [Proposition 6.22](#) gives $\Gamma \vdash \varphi \rightarrow \psi$. If $\psi \in \{\varphi\}$ then $\Gamma \vdash \varphi \rightarrow \psi$ because then last **sentence** $\varphi \rightarrow \psi$ is the same as $\varphi \rightarrow \varphi$, and we have **derived** that in [Example 6.10](#).

For the inductive step, suppose a **derivation** of ψ from $\Gamma \cup \{\varphi\}$ ends with a step ψ which is justified by modus ponens. (If it is not justified by modus ponens, $\psi \in \Gamma$, $\psi \equiv \varphi$, or ψ is an axiom, and the same reasoning as in the induction basis applies.) Then some previous steps in the **derivation** are $\chi \rightarrow \psi$ and χ , for some **formula** χ , i.e., $\Gamma \cup \{\varphi\} \vdash \chi \rightarrow \psi$ and $\Gamma \cup \{\varphi\} \vdash \chi$, and the respective derivations are shorter, so the inductive hypothesis applies to them. We thus have both:

$$\begin{aligned} \Gamma \vdash \varphi \rightarrow (\chi \rightarrow \psi); \\ \Gamma \vdash \varphi \rightarrow \chi. \end{aligned}$$

But also

$$\Gamma \vdash (\varphi \rightarrow (\chi \rightarrow \psi)) \rightarrow ((\varphi \rightarrow \chi) \rightarrow (\varphi \rightarrow \psi)),$$

by eq. (6.8), and two applications of Proposition 6.22 give $\Gamma \vdash \varphi \rightarrow \psi$, as required. \square

Notice how eq. (6.7) and eq. (6.8) were chosen precisely so that the Deduction Theorem would hold.

The following are some useful facts about *derivability*, which we leave as exercises.

Proposition 6.24.

*fol:axd:ded:
prop:derivfacts
fol:axd:ded:
derivfacts:a
fol:axd:ded:
derivfacts:b
fol:axd:ded:
derivfacts:c
fol:axd:ded:
derivfacts:d
fol:axd:ded:
derivfacts:e*

1. $\vdash (\varphi \rightarrow \psi) \rightarrow ((\psi \rightarrow \chi) \rightarrow (\varphi \rightarrow \chi))$;
2. If $\Gamma \cup \{\neg\varphi\} \vdash \neg\psi$ then $\Gamma \cup \{\psi\} \vdash \varphi$ (*Contraposition*);
3. $\{\varphi, \neg\varphi\} \vdash \psi$ (*Ex Falso Quodlibet, Explosion*);
4. $\{\neg\neg\varphi\} \vdash \varphi$ (*Double Negation Elimination*);
5. If $\Gamma \vdash \neg\neg\varphi$ then $\Gamma \vdash \varphi$;

Problem 6.3. Prove Proposition 6.24

6.8 The Deduction Theorem with Quantifiers

*fol:axd:ddq:
sec*

Theorem 6.25 (Deduction Theorem). *If $\Gamma \cup \{\varphi\} \vdash \psi$ and then $\Gamma \vdash \varphi \rightarrow \psi$.*

thm:deduction-thm-q

Proof. We again proceed by induction on the length of the *derivation* of ψ from $\Gamma \cup \{\varphi\}$.

The proof of the induction basis is identical to that in the proof of Theorem 6.23.

For the inductive step, suppose again that the *derivation* of ψ from $\Gamma \cup \{\varphi\}$ ends with a step ψ which is justified by an inference rule. If the inference rule is modus ponens, we proceed as in the proof of Theorem 6.23. If the inference rule is QR, we know that $\psi \equiv \chi \rightarrow \forall x \theta(x)$ and a *formula* of the form $\chi \rightarrow \theta(a)$ appears earlier in the *derivation*, where a does not occur in χ , φ , or Γ . We thus have that

$$\Gamma \cup \{\varphi\} \vdash \chi \rightarrow \theta(a)$$

and the induction hypothesis applies, i.e., we have that

$$\Gamma \vdash \varphi \rightarrow \theta(a)$$

By

$$\vdash (\varphi \rightarrow (\chi \rightarrow \theta(a))) \rightarrow ((\varphi \wedge \chi) \rightarrow \theta(a))$$

and modus ponens we get

$$\Gamma \vdash (\varphi \wedge \chi) \rightarrow \theta(a).$$

Since the eigenvariable condition still applies, we can add a step to this [derivation](#) justified by QR, and get:

$$\Gamma \vdash (\varphi \wedge \chi) \rightarrow \forall x \theta(x)$$

We also have

$$\vdash ((\varphi \wedge \chi) \rightarrow \forall x \theta(x)) \rightarrow (\varphi \rightarrow (\chi \rightarrow \forall x \theta(x)))$$

so by modus ponens,

$$\Gamma \vdash \varphi \rightarrow (\chi \rightarrow \forall x \theta(x))$$

i.e., $\Gamma \vdash \psi$.

We leave the case where ψ is justified by the rule QR, but is of the form $\exists x \theta(x) \rightarrow \chi$, as an exercise. \square

Problem 6.4. Complete the proof of [Theorem 6.25](#).

6.9 Derivability and Consistency

We will now establish a number of properties of the [derivability](#) relation. They are independently interesting, but each will play a role in the proof of the completeness theorem.

[fol:axd:prv:](#)
[sec](#)

Proposition 6.26. *If $\Gamma \vdash \varphi$ and $\Gamma \cup \{\varphi\}$ is inconsistent, then Γ is inconsistent.*

[fol:axd:prv:](#)
[prop:provability-contr](#)

Proof. If $\Gamma \cup \{\varphi\}$ is inconsistent, then $\Gamma \cup \{\varphi\} \vdash \perp$. By [Proposition 6.17](#), $\Gamma \vdash \psi$ for every $\psi \in \Gamma$. Since also $\Gamma \vdash \varphi$ by hypothesis, $\Gamma \vdash \psi$ for every $\psi \in \Gamma \cup \{\varphi\}$. By [Proposition 6.19](#), $\Gamma \vdash \perp$, i.e., Γ is inconsistent. \square

Proposition 6.27. *$\Gamma \vdash \varphi$ iff $\Gamma \cup \{\neg\varphi\}$ is inconsistent.*

[fol:axd:prv:](#)
[prop:prov-incons](#)

Proof. First suppose $\Gamma \vdash \varphi$. Then $\Gamma \cup \{\neg\varphi\} \vdash \varphi$ by [Proposition 6.18](#). $\Gamma \cup \{\neg\varphi\} \vdash \neg\varphi$ by [Proposition 6.17](#). We also have $\vdash \neg\varphi \rightarrow (\varphi \rightarrow \perp)$ by [eq. \(6.10\)](#). So by two applications of [Proposition 6.22](#), we have $\Gamma \cup \{\neg\varphi\} \vdash \perp$.

Now assume $\Gamma \cup \{\neg\varphi\}$ is inconsistent, i.e., $\Gamma \cup \{\neg\varphi\} \vdash \perp$. By the deduction theorem, $\Gamma \vdash \neg\varphi \rightarrow \perp$. $\Gamma \vdash (\neg\varphi \rightarrow \perp) \rightarrow \neg\neg\varphi$ by [eq. \(6.13\)](#), so $\Gamma \vdash \neg\neg\varphi$ by [Proposition 6.22](#). Since $\Gamma \vdash \neg\neg\varphi \rightarrow \varphi$ ([eq. \(6.14\)](#)), we have $\Gamma \vdash \varphi$ by [Proposition 6.22](#) again. \square

Problem 6.5. Prove that $\Gamma \vdash \neg\varphi$ iff $\Gamma \cup \{\varphi\}$ is inconsistent.

fol:axd:prv:
prop:explicit-inc

Proposition 6.28. *If $\Gamma \vdash \varphi$ and $\neg\varphi \in \Gamma$, then Γ is inconsistent.*

Proof. $\Gamma \vdash \neg\varphi \rightarrow (\varphi \rightarrow \perp)$ by eq. (6.10). $\Gamma \vdash \perp$ by two applications of Proposition 6.22. \square

fol:axd:prv:
prop:provability-exhaustive

Proposition 6.29. *If $\Gamma \cup \{\varphi\}$ and $\Gamma \cup \{\neg\varphi\}$ are both inconsistent, then Γ is inconsistent.*

Proof. Exercise. \square

Problem 6.6. Prove Proposition 6.29

6.10 Derivability and the Propositional Connectives

fol:axd:ppr:
sec

Proposition 6.30.

fol:axd:ppr:
prop:provability-land
fol:axd:ppr:
prop:provability-land-left
fol:axd:ppr:
prop:provability-land-right

1. Both $\varphi \wedge \psi \vdash \varphi$ and $\varphi \wedge \psi \vdash \psi$
2. $\varphi, \psi \vdash \varphi \wedge \psi$.

Proof. 1. From eq. (6.1) and eq. (6.1) by modus ponens.
2. From eq. (6.3) by two applications of modus ponens. \square

fol:axd:ppr:
prop:provability-lor

Proposition 6.31.

1. $\varphi \vee \psi, \neg\varphi, \neg\psi$ is inconsistent.
2. Both $\varphi \vdash \varphi \vee \psi$ and $\psi \vdash \varphi \vee \psi$.

Proof. 1. From eq. (6.9) we get $\vdash \neg\varphi \rightarrow (\varphi \rightarrow \perp)$ and $\vdash \neg\psi \rightarrow (\psi \rightarrow \perp)$. So by the deduction theorem, we have $\{\neg\varphi\} \vdash \varphi \rightarrow \perp$ and $\{\neg\psi\} \vdash \psi \rightarrow \perp$. From eq. (6.6) we get $\{\neg\varphi, \neg\psi\} \vdash (\varphi \vee \psi) \rightarrow \perp$. By the deduction theorem, $\{\varphi \vee \psi, \neg\varphi, \neg\psi\} \vdash \perp$.

2. From eq. (6.4) and eq. (6.5) by modus ponens. \square

fol:axd:ppr:
prop:provability-lif

Proposition 6.32.

fol:axd:ppr:
prop:provability-lif-left
fol:axd:ppr:
prop:provability-lif-right

1. $\varphi, \varphi \rightarrow \psi \vdash \psi$.
2. Both $\neg\varphi \vdash \varphi \rightarrow \psi$ and $\psi \vdash \varphi \rightarrow \psi$.

Proof. 1. We can derive:

1. φ HYP
2. $\varphi \rightarrow \psi$ HYP
3. ψ 1, 2, MP

2. By eq. (6.10) and eq. (6.7) and the deduction theorem, respectively. \square

6.11 Derivability and the Quantifiers

Theorem 6.33. *If c is a constant symbol not occurring in Γ or $\varphi(x)$ and $\Gamma \vdash \varphi(c)$, then $\Gamma \vdash \forall x \varphi(x)$.*

fol.axd:qpr:
sec
fol.axd:qpr:
thm:strong-generalization

Proof. By the deduction theorem, $\Gamma \vdash \top \rightarrow \varphi(c)$. Since c does not occur in Γ or \top , we get $\Gamma \vdash \top \rightarrow \varphi(c)$. By the deduction theorem again, $\Gamma \vdash \forall x \varphi(x)$. \square

Proposition 6.34.

fol.axd:qpr:
prop:provability-quantifiers

1. $\varphi(t) \vdash \exists x \varphi(x)$.
2. $\forall x \varphi(x) \vdash \varphi(t)$.

Proof. 1. By eq. (6.16) and the deduction theorem.

2. By eq. (6.15) and the deduction theorem. \square

6.12 Soundness

explanation

A derivation system, such as axiomatic deduction, is *sound* if it cannot derive things that do not actually hold. Soundness is thus a kind of guaranteed safety property for derivation systems. Depending on which proof theoretic property is in question, we would like to know for instance, that

fol.axd:sou:
sec

1. every derivable sentence is valid;
2. if a sentence is derivable from some others, it is also a consequence of them;
3. if a set of sentences is inconsistent, it is unsatisfiable.

These are important properties of a derivation system. If any of them do not hold, the derivation system is deficient—it would derive too much. Consequently, establishing the soundness of a derivation system is of the utmost importance.

Proposition 6.35. *If φ is an axiom, then $\mathfrak{M}, s \models \varphi$ for each structure \mathfrak{M} and assignment s .*

Proof. We first verify that all the axioms are valid. For instance, here is the case for eq. (6.15): suppose t is free for x in φ , and assume $\mathfrak{M}, s \models \forall x \varphi$. Then by definition of satisfaction, for each $s' \sim_x s$, also $\mathfrak{M}, s' \models \varphi$, and in particular this holds when $s'(x) = \text{Val}_s^{\mathfrak{M}}(t)$. By Proposition 1.45, $\mathfrak{M}, s \models \varphi[t/x]$. This shows that $\mathfrak{M}, s \models (\forall x \varphi \rightarrow \varphi[t/x])$. \square

fol:axd:sou:
thm:soundness

Theorem 6.36 (Soundness). *If $\Gamma \vdash \varphi$ then $\Gamma \models \varphi$.*

Proof. By induction on the length of the **derivation** of φ from Γ . If there are no steps justified by inferences, then all **formulas** in the derivation are either instances of axioms or are in Γ . By the previous proposition, all the axioms are valid, and hence if φ is an axiom then $\Gamma \models \varphi$. If $\varphi \in \Gamma$, then trivially $\Gamma \models \varphi$.

If the last step of the derivation of φ is justified by modus ponens, then there are **formulas** ψ and $\psi \rightarrow \varphi$ in the **derivation**, and the induction hypothesis applies to the part of the **derivation** ending in those **formulas** (since they contain at least one fewer steps justified by an inference). So, by induction hypothesis, $\Gamma \models \psi$ and $\Gamma \models \psi \rightarrow \varphi$. Then $\Gamma \models \varphi$ by [Theorem 1.52](#).

Now suppose the last step is justified by QR. Then that step has the form $\chi \rightarrow \forall x B(x)$ and there is a preceding step $\chi \rightarrow \psi(c)$ with c not in Γ , χ , or $\forall x B(x)$. By induction hypothesis, $\Gamma \models \chi \rightarrow \forall x B(x)$. By [Theorem 1.52](#), $\Gamma \cup \{\chi\} \models \psi(c)$.

Consider some structure \mathfrak{M} such that $\mathfrak{M} \models \Gamma \cup \{\chi\}$. We need to show that $\mathfrak{M} \models \forall x \psi(x)$. Since $\forall x \psi(x)$ is a **sentence**, this means we have to show that for every variable assignment s , $\mathfrak{M}, s \models \psi(x)$ ([Proposition 1.41](#)). Since $\Gamma \cup \{\chi\}$ consists entirely of sentences, $\mathfrak{M}, s \models \theta$ for all $\theta \in \Gamma$ by [Definition 1.34](#). Let \mathfrak{M}' be like \mathfrak{M} except that $c^{\mathfrak{M}'} = s(x)$. Since c does not occur in Γ or χ , $\mathfrak{M}' \models \Gamma \cup \{\chi\}$ by [Corollary 1.43](#). Since $\Gamma \cup \{\chi\} \models \psi(c)$, $\mathfrak{M}' \models \psi(c)$. Since $\psi(c)$ is a **sentence**, $\mathfrak{M}, s \models \psi(c)$ by [Proposition 1.40](#). $\mathfrak{M}', s \models \psi(x)$ iff $\mathfrak{M}' \models \psi(c)$ by [Proposition 1.45](#) (recall that $\psi(c)$ is just $\psi(x)[c/x]$). So, $\mathfrak{M}', s \models \psi(x)$. Since c does not occur in $\psi(x)$, by [Proposition 1.42](#), $\mathfrak{M}, s \models \psi(x)$. But s was an arbitrary variable assignment, so $\mathfrak{M} \models \forall x \psi(x)$. Thus $\Gamma \cup \{\chi\} \models \forall x \psi(x)$. By [Theorem 1.52](#), $\Gamma \models \chi \rightarrow \forall x \psi(x)$.

The case where φ is justified by QR but is of the form $\exists x \psi(x) \rightarrow \chi$ is left as an exercise. \square

Problem 6.7. Complete the proof of [Theorem 6.36](#).

fol:axd:sou:
cor:weak-soundness

Corollary 6.37. *If $\vdash \varphi$, then φ is valid.*

fol:axd:sou:
cor:consistency-soundness

Corollary 6.38. *If Γ is satisfiable, then it is consistent.*

Proof. We prove the contrapositive. Suppose that Γ is not consistent. Then $\Gamma \vdash \perp$, i.e., there is a **derivation** of \perp from Γ . By [Theorem 6.36](#), any **structure** \mathfrak{M} that satisfies Γ must satisfy \perp . Since $\mathfrak{M} \not\models \perp$ for every **structure** \mathfrak{M} , no \mathfrak{M} can satisfy Γ , i.e., Γ is not satisfiable. \square

6.13 Derivations with Identity predicate

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sec

In order to accommodate $=$ in **derivations**, we simply add new axiom schemas. The definition of **derivation** and \vdash remains the same, we just also allow the new axioms.

Definition 6.39 (Axioms for **identity predicate**).

$$t = t, \tag{6.17} \text{ fol:axd:ide:}$$

$$t_1 = t_2 \rightarrow (\psi(t_1) \rightarrow \psi(t_2)), \tag{6.18} \text{ fol:axd:ide:}$$

ax:iden2

for any ground terms t, t_1, t_2 .

Proposition 6.40. *The axioms eq. (6.17) and eq. (6.18) are valid.*

fol:axd:ide:
prop:sound

Proof. Exercise. □

Problem 6.8. Prove [Proposition 6.40](#).

Proposition 6.41. $\Gamma \vdash t = t$, for any term t and set Γ .

fol:axd:ide:
prop:iden1

Proposition 6.42. If $\Gamma \vdash \varphi(t_1)$ and $\Gamma \vdash t_1 = t_2$, then $\Gamma \vdash \varphi(t_2)$.

fol:axd:ide:
prop:iden2

Proof. The **formula**

$$(t_1 = t_2 \rightarrow (\varphi(t_1) \rightarrow \varphi(t_2)))$$

is an instance of [eq. \(6.18\)](#). The conclusion follows by two applications of MP. □

Chapter 7

The Completeness Theorem

7.1 Introduction

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sec

The completeness theorem is one of the most fundamental results about logic. It comes in two formulations, the equivalence of which we'll prove. In its first formulation it says something fundamental about the relationship between semantic consequence and our proof system: if a **sentence** φ follows from some **sentences** Γ , then there is also a **derivation** that establishes $\Gamma \vdash \varphi$. Thus, the proof system is as strong as it can possibly be without proving things that don't actually follow.

In its second formulation, it can be stated as a model existence result: every consistent set of **sentences** is satisfiable. Consistency is a proof-theoretic notion: it says that our proof system is unable to produce certain **derivations**. But who's to say that just because there are no **derivations** of a certain sort from Γ , it's guaranteed that there is a **structure** \mathfrak{M} with $\mathfrak{M} \models \Gamma$? Before the completeness theorem was first proved—in fact before we had the proof systems we now do—the great German mathematician David Hilbert held the view that consistency of mathematical theories guarantees the existence of the objects they are about. He put it as follows in a letter to Gottlob Frege:

If the arbitrarily given axioms do not contradict one another with all their consequences, then they are true and the things defined by the axioms exist. This is for me the criterion of truth and existence.

Frege vehemently disagreed. The second formulation of the completeness theorem shows that Hilbert was right in at least the sense that if the axioms are consistent, then *some* **structure** exists that makes them all true.

These aren't the only reasons the completeness theorem—or rather, its proof—is important. It has a number of important consequences, some of which we'll discuss separately. For instance, since any **derivation** that shows $\Gamma \vdash \varphi$ is finite and so can only use finitely many of the **sentences** in Γ , it follows by the completeness theorem that if φ is a consequence of Γ , it is already a

consequence of a finite subset of Γ . This is called *compactness*. Equivalently, if every finite subset of Γ is consistent, then Γ itself must be consistent.

Although the compactness theorem follows from the completeness theorem via the detour through *derivations*, it is also possible to use the *the proof of the completeness theorem* to establish it directly. For what the proof does is take a set of *sentences* with a certain property—consistency—and constructs a *structure* out of this set that has certain properties (in this case, that it satisfies the set). Almost the very same construction can be used to directly establish compactness, by starting from “finitely satisfiable” sets of *sentences* instead of consistent ones. The construction also yields other consequences, e.g., that any satisfiable set of *sentences* has a finite or *denumerable* model. (This result is called the Löwenheim-Skolem theorem.) In general, the construction of *structures* from sets of *sentences* is used often in logic, and sometimes even in philosophy.

7.2 Outline of the Proof

The proof of the completeness theorem is a bit complex, and upon first reading it, it is easy to get lost. So let us outline the proof. The first step is a shift of perspective, that allows us to see a route to a proof. When completeness is thought of as “whenever $\Gamma \models \varphi$ then $\Gamma \vdash \varphi$,” it may be hard to even come up with an idea: for to show that $\Gamma \vdash \varphi$ we have to find a *derivation*, and it does not look like the hypothesis that $\Gamma \models \varphi$ helps us for this in any way. For some proof systems it is possible to directly construct a *derivation*, but we will take a slightly different tack. The shift in perspective required is this: completeness can also be formulated as: “if Γ is consistent, it has a model.” Perhaps we can use the information in Γ together with the hypothesis that it is consistent to construct a model. After all, we know what kind of model we are looking for: one that is as Γ describes it!

fol:com:out:
sec

If Γ contains only atomic sentences, it is easy to construct a model for it: for atomic sentences are all of the form $P(a_1, \dots, a_n)$ where the a_i are *constant symbols*. So all we have to do is come up with a *domain* $|\mathfrak{M}|$ and an interpretation for P so that $\mathfrak{M} \models P(a_1, \dots, a_n)$. But nothing’s easier than that: put $|\mathfrak{M}| = \mathbb{N}$, $c_i^{\mathfrak{M}} = i$, and for every $P(a_1, \dots, a_n) \in \Gamma$, put the tuple $\langle k_1, \dots, k_n \rangle$ into $P^{\mathfrak{M}}$, where k_i is the index of the constant symbol a_i (i.e., $a_i \equiv c_{k_i}$).

Now suppose Γ contains some sentence $\neg\psi$, with ψ atomic. We might worry that the construction of \mathfrak{M} interferes with the possibility of making $\neg\psi$ true. But here’s where the consistency of Γ comes in: if $\neg\psi \in \Gamma$, then $\psi \notin \Gamma$, or else Γ would be inconsistent. And if $\psi \notin \Gamma$, then according to our construction of \mathfrak{M} , $\mathfrak{M} \not\models \psi$, so $\mathfrak{M} \models \neg\psi$. So far so good.

Now what if Γ contains complex, non-atomic formulas? Say, it contains $\varphi \wedge \psi$. Then we should proceed as if both φ and ψ were in Γ . And if $\varphi \vee \psi \in \Gamma$, then we will have to make at least one of them true, i.e., proceed as if one of them was in Γ .

This suggests the following idea: we add additional sentences to Γ so as to (a) keep the resulting set consistent and (b) make sure that for every possible atomic sentence φ , either φ is in the resulting set, or $\neg\varphi$, and (c) such that, whenever $\varphi \wedge \psi$ is in the set, so are both φ and ψ , if $\varphi \vee \psi$ is in the set, at least one of φ or ψ is also, etc. We keep doing this (potentially forever). Call the set of all sentences so added Γ^* . Then our construction above would provide us with a structure for which we could prove, by induction, that all sentences in Γ^* are true in \mathfrak{M} , and hence also all sentence in Γ since $\Gamma \subseteq \Gamma^*$. It turns out that guaranteeing (a) is enough. A set of sentences for which (a) holds is called *complete*. So our task will be to extend the consistent set Γ to a consistent and complete set Γ^* .

There is one wrinkle in this plan: if $\exists x \varphi(x) \in \Gamma$ we would hope to be able to pick some **constant symbol** c and add $\varphi(c)$ in this process. But how do we know we can always do that? Perhaps we only have a few **constant symbols** in our language, and for each one of them we have $\neg\varphi(c) \in \Gamma$. We can't also add $\varphi(c)$, since this would make the set inconsistent, and we wouldn't know whether \mathfrak{M} has to make $\varphi(c)$ or $\neg\varphi(c)$ true. Moreover, it might happen that Γ contains only sentences in a language that has no constant symbols at all (e.g., the language of set theory).

The solution to this problem is to simply add infinitely many constants at the beginning, plus sentences that connect them with the quantifiers in the right way. (Of course, we have to verify that this cannot introduce an inconsistency.)

Our original construction works well if we only have **constant symbols** in the atomic sentences. But the language might also contain **function symbols**. In that case, it might be tricky to find the right functions on \mathbb{N} to assign to these **function symbols** to make everything work. So here's another trick: instead of using i to interpret c_i , just take the set of **constant symbols** itself as the domain. Then \mathfrak{M} can assign every **constant symbol** to itself: $c_i^{\mathfrak{M}} = c_i$. But why not go all the way: let $|\mathfrak{M}|$ be all *terms* of the language! If we do this, there is an obvious assignment of functions (that take terms as arguments and have terms as values) to **function symbols**: we assign to the **function symbol** f_i^n the function which, given n terms t_1, \dots, t_n as input, produces the term $f_i^n(t_1, \dots, t_n)$ as value.

The last piece of the puzzle is what to do with $=$. The **predicate symbol** $=$ has a fixed interpretation: $\mathfrak{M} \models t = t'$ iff $\text{Val}^{\mathfrak{M}}(t) = \text{Val}^{\mathfrak{M}}(t')$. Now if we set things up so that the **value** of a term t is t itself, then this **structure** will make *no* sentence of the form $t = t'$ true unless t and t' are one and the same term. And of course this is a problem, since basically every interesting theory in a language with **function symbols** will have as theorems sentences $t = t'$ where t and t' are not the same term (e.g., in theories of arithmetic: $(0 + 0) = 0$). To solve this problem, we change the domain of \mathfrak{M} : instead of using terms as the objects in $|\mathfrak{M}|$, we use sets of terms, and each set is so that it contains all those terms which the sentences in Γ require to be equal. So, e.g., if Γ is a theory of arithmetic, one of these sets will contain: $0, (0 + 0), (0 \times 0)$, etc. This will be the set we assign to 0 , and it will turn out that this set is also the value of all the terms in it, e.g., also of $(0 + 0)$. Therefore, the sentence $(0 + 0) = 0$ will

be true in this revised [structure](#).

So here's what we'll do. First we investigate the properties of [complete](#) consistent sets, in particular we prove that a [complete](#) consistent set contains $\varphi \wedge \psi$ iff it contains both φ and ψ , $\varphi \vee \psi$ iff it contains at least one of them, etc. ([Proposition 7.2](#)). Then we define and investigate “saturated” sets of sentences. A saturated set is one which contains conditionals that link each quantified [sentence](#) to instances of it ([Definition 7.5](#)). We show that any consistent set Γ can always be extended to a saturated set Γ' ([Lemma 7.6](#)). If a set is consistent, saturated, and [complete](#) it also has the property that it contains $\exists x \varphi(x)$ iff it contains $\varphi(t)$ for some closed term t and $\forall x \varphi(x)$ iff it contains $\varphi(t)$ for all closed terms t ([Proposition 7.7](#)). We'll then take the saturated consistent set Γ' and show that it can be extended to a saturated, consistent, and [complete](#) set Γ^* ([Lemma 7.8](#)). This set Γ^* is what we'll use to define our term model $\mathfrak{M}(\Gamma^*)$. The term model has the set of closed terms as its domain, and the interpretation of its [predicate symbols](#) is given by the atomic [sentences](#) in Γ^* ([Definition 7.9](#)). We'll use the properties of consistent, complete, saturated sets to show that indeed $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\varphi \in \Gamma^*$ ([Lemma 7.11](#)), and thus in particular, $\mathfrak{M}(\Gamma^*) \models \Gamma$. Finally, we'll consider how to define a term model if Γ contains = as well ([Definition 7.15](#)) and show that it satisfies Γ^* ([Lemma 7.17](#)).

7.3 Complete Consistent Sets of Sentences

Definition 7.1 (Complete set). A set Γ of [sentences](#) is [complete](#) iff for any [sentence](#) φ , either $\varphi \in \Gamma$ or $\neg\varphi \in \Gamma$.

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sec

fol:com:ccs:
def:complete-set

[explanation](#) [Complete](#) sets of sentences leave no questions unanswered. For any [sentence](#) A , Γ “says” if φ is true or false. The importance of [complete](#) sets extends beyond the proof of the completeness theorem. A theory which is [complete](#) and axiomatizable, for instance, is always decidable.

[explanation](#) [Complete](#) consistent sets are important in the completeness proof since we can guarantee that every consistent set of [sentences](#) Γ is contained in a [complete](#) consistent set Γ^* . A [complete](#) consistent set contains, for each [sentence](#) φ , either φ or its negation $\neg\varphi$, but not both. This is true in particular for atomic [sentences](#), so from a [complete](#) consistent set in a language suitably expanded by [constant symbols](#), we can construct a [structure](#) where the interpretation of [predicate symbols](#) is defined according to which atomic [sentences](#) are in Γ^* . This [structure](#) can then be shown to make all [sentences](#) in Γ^* (and hence also all those in Γ) true. The proof of this latter fact requires that $\neg\varphi \in \Gamma^*$ iff $\varphi \notin \Gamma^*$, $(\varphi \vee \psi) \in \Gamma^*$ iff $\varphi \in \Gamma^*$ or $\psi \in \Gamma^*$, etc.

In what follows, we will often tacitly use the properties of reflexivity, monotonicity, and transitivity of \vdash (see [sections 4.8, 5.7 and 6.6](#)).

Proposition 7.2. *Suppose Γ is [complete](#) and consistent. Then:*

fol:com:ccs:
prop:ccs

1. If $\Gamma \vdash \varphi$, then $\varphi \in \Gamma$.

fol:com:ccs:
prop:ccs-prov-in

*fol.com.ccs:
prop.ccs-and*

2. $\varphi \wedge \psi \in \Gamma$ iff both $\varphi \in \Gamma$ and $\psi \in \Gamma$.

*fol.com.ccs:
prop.ccs-or*

3. $\varphi \vee \psi \in \Gamma$ iff either $\varphi \in \Gamma$ or $\psi \in \Gamma$.

*fol.com.ccs:
prop.ccs-if*

4. $\varphi \rightarrow \psi \in \Gamma$ iff either $\varphi \notin \Gamma$ or $\psi \in \Gamma$.

Proof. Let us suppose for all of the following that Γ is **complete** and consistent.

1. If $\Gamma \vdash \varphi$, then $\varphi \in \Gamma$.

Suppose that $\Gamma \vdash \varphi$. Suppose to the contrary that $\varphi \notin \Gamma$. Since Γ is **complete**, $\neg\varphi \in \Gamma$. By [Propositions 4.20, 5.20](#) and [6.28](#), Γ is inconsistent. This contradicts the assumption that Γ is consistent. Hence, it cannot be the case that $\varphi \notin \Gamma$, so $\varphi \in \Gamma$.

2. $\varphi \wedge \psi \in \Gamma$ iff both $\varphi \in \Gamma$ and $\psi \in \Gamma$:

For the forward direction, suppose $\varphi \wedge \psi \in \Gamma$. Then by [Propositions 4.22, 5.22](#) and [6.30](#), item (1), $\Gamma \vdash \varphi$ and $\Gamma \vdash \psi$. By (1), $\varphi \in \Gamma$ and $\psi \in \Gamma$, as required.

For the reverse direction, let $\varphi \in \Gamma$ and $\psi \in \Gamma$. By [Propositions 4.22, 5.22](#) and [6.30](#), item (2), $\Gamma \vdash \varphi \wedge \psi$. By (1), $\varphi \wedge \psi \in \Gamma$.

3. First we show that if $\varphi \vee \psi \in \Gamma$, then either $\varphi \in \Gamma$ or $\psi \in \Gamma$. Suppose $\varphi \vee \psi \in \Gamma$ but $\varphi \notin \Gamma$ and $\psi \notin \Gamma$. Since Γ is **complete**, $\neg\varphi \in \Gamma$ and $\neg\psi \in \Gamma$. By [Propositions 4.23, 5.23](#) and [6.31](#), item (1), Γ is inconsistent, a contradiction. Hence, either $\varphi \in \Gamma$ or $\psi \in \Gamma$.

For the reverse direction, suppose that $\varphi \in \Gamma$ or $\psi \in \Gamma$. By [Propositions 4.23, 5.23](#) and [6.31](#), item (2), $\Gamma \vdash \varphi \vee \psi$. By (1), $\varphi \vee \psi \in \Gamma$, as required.

4. For the forward direction, suppose $\varphi \rightarrow \psi \in \Gamma$, and suppose to the contrary that $\varphi \in \Gamma$ and $\psi \notin \Gamma$. On these assumptions, $\varphi \rightarrow \psi \in \Gamma$ and $\varphi \in \Gamma$. By [Propositions 4.24, 5.24](#) and [6.32](#), item (1), $\Gamma \vdash \psi$. But then by (1), $\psi \in \Gamma$, contradicting the assumption that $\psi \notin \Gamma$.

For the reverse direction, first consider the case where $\varphi \notin \Gamma$. Since Γ is **complete**, $\neg\varphi \in \Gamma$. By [Propositions 4.24, 5.24](#) and [6.32](#), item (2), $\Gamma \vdash \varphi \rightarrow \psi$. Again by (1), we get that $\varphi \rightarrow \psi \in \Gamma$, as required.

Now consider the case where $\psi \in \Gamma$. By [Propositions 4.24, 5.24](#) and [6.32](#), item (2) again, $\Gamma \vdash \varphi \rightarrow \psi$. By (1), $\varphi \rightarrow \psi \in \Gamma$.

□

Problem 7.1. Complete the proof of [Proposition 7.2](#).

7.4 Henkin Expansion

explanation

Part of the challenge in proving the completeness theorem is that the model we construct from a complete consistent set Γ must make all the quantified formulas in Γ true. In order to guarantee this, we use a trick due to Leon Henkin. In essence, the trick consists in expanding the language by infinitely many constant symbols and adding, for each formula with one free variable $\varphi(x)$ a formula of the form $\exists x \varphi \rightarrow \varphi(c)$, where c is one of the new constant symbols. When we construct the structure satisfying Γ , this will guarantee that each true existential sentence has a witness among the new constants.

fol:com:hen:
sec

Proposition 7.3. *If Γ is consistent in \mathcal{L} and \mathcal{L}' is obtained from \mathcal{L} by adding a denumerable set of new constant symbols d_0, d_1, \dots , then Γ is consistent in \mathcal{L}' .*

fol:com:hen:
prop:lang-exp

Definition 7.4 (Saturated set). A set Γ of formulas of a language \mathcal{L} is *saturated* iff for each formula $\varphi(x) \in \text{Frm}(\mathcal{L})$ with one free variable x there is a constant symbol $c \in \mathcal{L}$ such that $\exists x \varphi(x) \rightarrow \varphi(c) \in \Gamma$.

The following definition will be used in the proof of the next theorem.

Definition 7.5. Let \mathcal{L}' be as in Proposition 7.3. Fix an enumeration $\varphi_0(x_0), \varphi_1(x_1), \dots$ of all formulas $\varphi_i(x_i)$ of \mathcal{L}' in which one variable (x_i) occurs free. We define the sentences θ_n by induction on n .

fol:com:hen:
defn:henkin-exp

Let c_0 be the first constant symbol among the d_i we added to \mathcal{L} which does not occur in $\varphi_0(x_0)$. Assuming that $\theta_0, \dots, \theta_{n-1}$ have already been defined, let c_n be the first among the new constant symbols d_i that occurs neither in $\theta_0, \dots, \theta_{n-1}$ nor in $\varphi_n(x_n)$.

Now let θ_n be the formula $\exists x_n \varphi_n(x_n) \rightarrow \varphi_n(c_n)$.

Lemma 7.6. *Every consistent set Γ can be extended to a saturated consistent set Γ' .*

fol:com:hen:
lem:henkin

Proof. Given a consistent set of sentences Γ in a language \mathcal{L} , expand the language by adding a denumerable set of new constant symbols to form \mathcal{L}' . By Proposition 7.3, Γ is still consistent in the richer language. Further, let θ_i be as in Definition 7.5. Let

$$\begin{aligned} \Gamma_0 &= \Gamma \\ \Gamma_{n+1} &= \Gamma_n \cup \{\theta_n\} \end{aligned}$$

i.e., $\Gamma_{n+1} = \Gamma \cup \{\theta_0, \dots, \theta_n\}$, and let $\Gamma' = \bigcup_n \Gamma_n$. Γ' is clearly saturated.

If Γ' were inconsistent, then for some n , Γ_n would be inconsistent (Exercise: explain why). So to show that Γ' is consistent it suffices to show, by induction on n , that each set Γ_n is consistent.

The induction basis is simply the claim that $\Gamma_0 = \Gamma$ is consistent, which is the hypothesis of the theorem. For the induction step, suppose that Γ_n is consistent but $\Gamma_{n+1} = \Gamma_n \cup \{\theta_n\}$ is inconsistent. Recall that θ_n is $\exists x_n \varphi_n(x_n) \rightarrow \varphi_n(c_n) \rightarrow$

$\varphi_n(c_n)$, where $\varphi_n(x_n)$ is a **formula** of \mathcal{L}' with only the variable x_n free. By the way we've chosen the c_n (see [Definition 7.5](#)), c_n does not occur in $A_n(x_n)$ nor in Γ_n .

If $\Gamma_n \cup \{\theta_n\}$ is inconsistent, then $\Gamma_n \vdash \neg\theta_n$, and hence both of the following hold:

$$\Gamma_n \vdash \exists x_n \varphi_n(x_n) \quad \Gamma_n \vdash \neg\varphi_n(c_n)$$

Since c_n does not occur in Γ_n or in $\varphi_n(x_n)$, [Theorems 4.25, 5.25 and 6.33](#) applies. From $\Gamma_n \vdash \neg\varphi_n(c_n)$, we obtain $\Gamma_n \vdash \forall x_n \neg\varphi_n(x_n)$. Thus we have that both $\Gamma_n \vdash \exists x_n \varphi_n$ and $\Gamma_n \vdash \forall x_n \neg\varphi_n(x_n)$, so Γ_n itself is inconsistent. (Note that $\forall x_n \neg\varphi_n(x_n) \vdash \neg\exists x_n \varphi_n(x_n)$.) Contradiction: Γ_n was supposed to be consistent. Hence $\Gamma_n \cup \{\theta_n\}$ is consistent. \square

We'll now show that *complete*, consistent sets which are saturated have the property that it contains a universally quantified **sentence** iff it contains all its instances and it contains an existentially quantified **sentence** iff it contains at least one instance. We'll use this to show that the **structure** we'll generate from a complete, consistent, saturated set makes all its quantified sentences true.

[explanation](#)

Proposition 7.7. *Suppose Γ is complete, consistent, and saturated.*

[fol:com:hen:](#)
[prop:saturated-instances](#)

1. $\exists x \varphi(x) \in \Gamma$ iff $\varphi(t) \in \Gamma$ for at least one closed term t .
2. $\forall x \varphi(x) \in \Gamma$ iff $\varphi(t) \in \Gamma$ for all closed terms t .

Proof. 1. First suppose that $\exists x \varphi(x) \in \Gamma$. Because Γ is saturated, $(\exists x \varphi(x) \rightarrow \varphi(c)) \in \Gamma$ for some **constant symbol** c . By [Propositions 4.24, 5.24 and 6.32](#), item (1), and [Proposition 7.2\(1\)](#), $\varphi(c) \in \Gamma$.

For the other direction, saturation is not necessary: Suppose $\varphi(t) \in \Gamma$. Then $\Gamma \vdash \exists x \varphi(x)$ by [Propositions 4.26, 5.26 and 6.34](#), item (1). By [Proposition 7.2\(1\)](#), $\exists x \varphi(x) \in \Gamma$.

2. Suppose that $\varphi(t) \in \Gamma$ for all closed terms t . By way of contradiction, assume $\forall x \varphi(x) \notin \Gamma$. Since Γ is complete, $\neg\forall x \varphi(x) \in \Gamma$. By saturation, $(\exists x \neg\varphi(x) \rightarrow \neg\varphi(c)) \in \Gamma$ for some **constant symbol** c . By assumption, since c is a closed term, $\varphi(c) \in \Gamma$. But this would make Γ inconsistent. (Exercise: give the derivation that shows

$$\neg\forall x \varphi(x), \exists x \neg\varphi(x) \rightarrow \neg\varphi(c), \varphi(c)$$

is inconsistent.)

For the reverse direction, we do not need saturation: Suppose $\forall x \varphi(x) \in \Gamma$. Then $\Gamma \vdash \varphi(t)$ by [Propositions 4.26, 5.26 and 6.34](#), item (2). We get $\varphi(t) \in \Gamma$ by [Proposition 7.2](#).

\square

7.5 Lindenbaum's Lemma

explanation

We now prove a lemma that shows that any consistent set of **sentences** is contained in some set of sentences which is not just consistent, but also **complete**. The proof works by adding one sentence at a time, guaranteeing at each step that the set remains consistent. We do this so that for every φ , either φ or $\neg\varphi$ gets added at some stage. The union of all stages in that construction then contains either φ or its negation $\neg\varphi$ and is thus complete. It is also consistent, since we made sure at each stage not to introduce an inconsistency.

fol:com:lin:
sec

Lemma 7.8 (Lindenbaum's Lemma). *Every consistent set Γ' in a language \mathcal{L}' can be extended to a **complete** and consistent set Γ^* .*

fol:com:lin:
lem:lindenbaum

Proof. Let Γ' be consistent. Let $\varphi_0, \varphi_1, \dots$ be an enumeration of all the **formulas** of \mathcal{L}' . Define $\Gamma_0 = \Gamma'$, and

$$\Gamma_{n+1} = \begin{cases} \Gamma_n \cup \{\varphi_n\} & \text{if } \Gamma_n \cup \{\varphi_n\} \text{ is consistent;} \\ \Gamma_n \cup \{\neg\varphi_n\} & \text{otherwise.} \end{cases}$$

Let $\Gamma^* = \bigcup_{n \geq 0} \Gamma_n$.

Each Γ_n is consistent: Γ_0 is consistent by definition. If $\Gamma_{n+1} = \Gamma_n \cup \{\varphi_n\}$, this is because the latter is consistent. If it isn't, $\Gamma_{n+1} = \Gamma_n \cup \{\neg\varphi_n\}$. We have to verify that $\Gamma_n \cup \{\neg\varphi_n\}$ is consistent. Suppose it's not. Then *both* $\Gamma_n \cup \{\varphi_n\}$ and $\Gamma_n \cup \{\neg\varphi_n\}$ are inconsistent. This means that Γ_n would be inconsistent by **Propositions 4.20, 5.20 and 6.28**, contrary to the induction hypothesis.

Every finite subset of Γ^* is a subset of Γ_n for some n , since each $\psi \in \Gamma^*$ not already in Γ' is added at some stage i . If n is the last one of these, then all ψ in the finite subset are in Γ_n . So, every finite subset of Γ^* is consistent. By **Propositions 4.17, 5.17 and 6.21**, Γ^* is consistent.

Every **sentence** of $\text{Frm}(\mathcal{L}')$ appears on the list used to define Γ^* . If $\varphi_n \notin \Gamma^*$, then that is because $\Gamma_n \cup \{\varphi_n\}$ was inconsistent. But then $\neg\varphi_n \in \Gamma^*$, so Γ^* is **complete**. \square

7.6 Construction of a Model

explanation

Right now we are not concerned about $=$, i.e., we only want to show that a consistent set Γ of **sentences** not containing $=$ is satisfiable. We first extend Γ to a consistent, **complete**, and saturated set Γ^* . In this case, the definition of a model $\mathfrak{M}(\Gamma^*)$ is simple: We take the set of closed terms of \mathcal{L}' as the domain. We assign every **constant symbol** to itself, and make sure that more generally, for every closed term t , $\text{Val}^{\mathfrak{M}(\Gamma^*)}(t) = t$. The **predicate symbols** are assigned extensions in such a way that an atomic **sentence** is true in $\mathfrak{M}(\Gamma^*)$ iff it is in Γ^* . This will obviously make all the atomic **sentences** in Γ^* true in $\mathfrak{M}(\Gamma^*)$. The rest are true provided the Γ^* we start with is consistent, complete, and saturated.

fol:com:mod:
sec

fol:com:mod:
defn:termmodel

Definition 7.9 (Term model). Let Γ^* be a **complete** and consistent, saturated set of **sentences** in a language \mathcal{L} . The *term model* $\mathfrak{M}(\Gamma^*)$ of Γ^* is the **structure** defined as follows:

1. The **domain** $|\mathfrak{M}(\Gamma^*)|$ is the set of all closed terms of \mathcal{L} .
2. The interpretation of a **constant symbol** c is c itself: $c^{\mathfrak{M}(\Gamma^*)} = c$.
3. The **function symbol** f is assigned the function which, given as arguments the closed terms t_1, \dots, t_n , has as value the closed term $f(t_1, \dots, t_n)$:

$$f^{\mathfrak{M}(\Gamma^*)}(t_1, \dots, t_n) = f(t_1, \dots, t_n)$$

4. If R is an n -place **predicate symbol**, then

$$\langle t_1, \dots, t_n \rangle \in R^{\mathfrak{M}(\Gamma^*)} \text{ iff } R(t_1, \dots, t_n) \in \Gamma^*.$$

A **structure** \mathfrak{M} may make an existentially quantified **sentence** $\exists x \varphi(x)$ true without there being an instance $\varphi(t)$ that it makes true. A **structure** \mathfrak{M} may make all instances $\varphi(t)$ of a universally quantified **sentence** $\forall x \varphi(x)$ true, without making $\forall x \varphi(x)$ true. This is because in general not every **element** of $|\mathfrak{M}|$ is the value of a closed term (\mathfrak{M} may not be covered). This is the reason the satisfaction relation is defined via variable assignments. However, for our term model $\mathfrak{M}(\Gamma^*)$ this wouldn't be necessary—because it is covered. This is the content of the next result. explanation

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prop:quant-termmodel

Proposition 7.10. *Let $\mathfrak{M}(\Gamma^*)$ be the term model of Definition 7.9.*

1. $\mathfrak{M}(\Gamma^*) \models \exists x \varphi(x)$ iff $\mathfrak{M} \models \varphi(t)$ for at least one term t .
2. $\mathfrak{M}(\Gamma^*) \models \forall x \varphi(x)$ iff $\mathfrak{M} \models \varphi(t)$ for all terms t .

Proof. 1. By Proposition 1.41, $\mathfrak{M}(\Gamma^*) \models \exists x \varphi(x)$ iff for at least one variable assignment s , $\mathfrak{M}(\Gamma^*), s \models \varphi(x)$. As $|\mathfrak{M}(\Gamma^*)|$ consists of the closed terms of \mathcal{L} , this is the case iff there is at least one closed term t such that $s(x) = t$ and $\mathfrak{M}(\Gamma^*), s \models \varphi(x)$. By Proposition 1.45, $\mathfrak{M}(\Gamma^*), s \models \varphi(x)$ iff $\mathfrak{M}(\Gamma^*), s \models \varphi(t)$, where $s(x) = t$. By Proposition 1.40, $\mathfrak{M}(\Gamma^*), s \models \varphi(t)$ iff $\mathfrak{M}(\Gamma^*) \models \varphi(t)$, since $\varphi(t)$ is a sentence.

2. By Proposition 1.41, $\mathfrak{M}(\Gamma^*) \models \forall x \varphi(x)$ iff for every variable assignment s , $\mathfrak{M}(\Gamma^*), s \models \varphi(x)$. Recall that $|\mathfrak{M}(\Gamma^*)|$ consists of the closed terms of \mathcal{L} , so for every closed term t , $s(x) = t$ is such a variable assignment, and for any variable assignment, $s(x)$ is some closed term t . By Proposition 1.45, $\mathfrak{M}(\Gamma^*), s \models \varphi(x)$ iff $\mathfrak{M}(\Gamma^*), s \models \varphi(t)$, where $s(x) = t$. By Proposition 1.40, $\mathfrak{M}(\Gamma^*), s \models \varphi(t)$ iff $\mathfrak{M}(\Gamma^*) \models \varphi(t)$, since $\varphi(t)$ is a sentence. □

Problem 7.2. Complete the proof of Proposition 7.10.

Lemma 7.11 (Truth Lemma). *Suppose φ does not contain $=$. Then $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\varphi \in \Gamma^*$.* fol.com.mod:
lem.truth

Proof. We prove both directions simultaneously, and by induction on φ .

1. $\varphi \equiv \perp$: $\mathfrak{M}(\Gamma^*) \not\models \perp$ by definition of satisfaction. On the other hand, $\perp \notin \Gamma^*$ since Γ^* is consistent.
2. $\varphi \equiv \top$: $\mathfrak{M}(\Gamma^*) \models \top$ by definition of satisfaction. On the other hand, $\top \in \Gamma^*$ since Γ^* is consistent and **complete**, and $\Gamma^* \vdash \top$.
3. $\varphi \equiv R(t_1, \dots, t_n)$: $\mathfrak{M}(\Gamma^*) \models R(t_1, \dots, t_n)$ iff $\langle t_1, \dots, t_n \rangle \in R^{\mathfrak{M}(\Gamma^*)}$ (by the definition of satisfaction) iff $R(t_1, \dots, t_n) \in \Gamma^*$ (by the construction of $\mathfrak{M}(\Gamma^*)$).
4. $\varphi \equiv \neg\psi$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\mathfrak{M}(\Gamma^*) \not\models \psi$ (by definition of satisfaction). By induction hypothesis, $\mathfrak{M}(\Gamma^*) \not\models \psi$ iff $\psi \notin \Gamma^*$. Since Γ^* is consistent and **complete**, $\psi \notin \Gamma^*$ iff $\neg\psi \in \Gamma^*$.
5. $\varphi \equiv \psi \wedge \chi$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff we have both $\mathfrak{M}(\Gamma^*) \models \psi$ and $\mathfrak{M}(\Gamma^*) \models \chi$ (by definition of satisfaction) iff both $\psi \in \Gamma^*$ and $\chi \in \Gamma^*$ (by the induction hypothesis). By [Proposition 7.2\(2\)](#), this is the case iff $(\psi \wedge \chi) \in \Gamma^*$.
6. $\varphi \equiv \psi \vee \chi$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff at $\mathfrak{M}(\Gamma^*) \models \psi$ or $\mathfrak{M}(\Gamma^*) \models \chi$ (by definition of satisfaction) iff $\psi \in \Gamma^*$ or $\chi \in \Gamma^*$ (by induction hypothesis). This is the case iff $(\psi \vee \chi) \in \Gamma^*$ (by [Proposition 7.2\(3\)](#)).
7. $\varphi \equiv \psi \rightarrow \chi$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\mathfrak{M}(\Gamma^*) \not\models \psi$ or $\mathfrak{M}(\Gamma^*) \models \chi$ (by definition of satisfaction) iff $\psi \notin \Gamma^*$ or $\chi \in \Gamma^*$ (by induction hypothesis). This is the case iff $(\psi \rightarrow \chi) \in \Gamma^*$ (by [Proposition 7.2\(4\)](#)).
8. $\varphi \equiv \forall x \psi(x)$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\mathfrak{M}(\Gamma^*) \models \psi(t)$ for all terms t ([Proposition 7.10](#)). By induction hypothesis, this is the case iff $\psi(t) \in \Gamma^*$ for all terms t , by [Proposition 7.7](#), this in turn is the case iff $\forall x \psi(x) \in \Gamma^*$.
9. $\varphi \equiv \exists x \psi(x)$: $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\mathfrak{M}(\Gamma^*) \models \psi(t)$ for at least one term t ([Proposition 7.10](#)). By induction hypothesis, this is the case iff $\psi(t) \in \Gamma^*$ for at least one term t . By [Proposition 7.7](#), this in turn is the case iff $\exists x \psi(x) \in \Gamma^*$.

□

Problem 7.3. Complete the proof of [Lemma 7.11](#).

7.7 Identity

fol:com:ide: sec The construction of the term model given in the preceding section is enough explanation to establish completeness for first-order logic for sets Γ that do not contain $=$. The term model satisfies every $\varphi \in \Gamma^*$ which does not contain $=$ (and hence all $\varphi \in \Gamma$). It does not work, however, if $=$ is present. The reason is that Γ^* then may contain a sentence $t = t'$, but in the term model the value of any term is that term itself. Hence, if t and t' are different terms, their values in the term model—i.e., t and t' , respectively—are different, and so $t = t'$ is false. We can fix this, however, using a construction known as “factoring.”

Definition 7.12. Let Γ^* be a consistent and complete set of sentences in \mathcal{L} . We define the relation \approx on the set of closed terms of \mathcal{L} by

$$t \approx t' \quad \text{iff} \quad t = t' \in \Gamma^*$$

fol:com:ide: prop:approx-equiv **Proposition 7.13.** *The relation \approx has the following properties:*

1. \approx is reflexive.
2. \approx is symmetric.
3. \approx is transitive.
4. If $t \approx t'$, f is a function symbol, and $t_1, \dots, t_{i-1}, t_{i+1}, \dots, t_n$ are terms, then

$$f(t_1, \dots, t_{i-1}, t, t_{i+1}, \dots, t_n) \approx f(t_1, \dots, t_{i-1}, t', t_{i+1}, \dots, t_n).$$

5. If $t \approx t'$, R is a predicate symbol, and $t_1, \dots, t_{i-1}, t_{i+1}, \dots, t_n$ are terms, then

$$R(t_1, \dots, t_{i-1}, t, t_{i+1}, \dots, t_n) \in \Gamma^* \quad \text{iff} \quad R(t_1, \dots, t_{i-1}, t', t_{i+1}, \dots, t_n) \in \Gamma^*.$$

Proof. Since Γ^* is consistent and complete, $t = t' \in \Gamma^*$ iff $\Gamma^* \vdash t = t'$. Thus it is enough to show the following:

1. $\Gamma^* \vdash t = t$ for all terms t .
2. If $\Gamma^* \vdash t = t'$ then $\Gamma^* \vdash t' = t$.
3. If $\Gamma^* \vdash t = t'$ and $\Gamma^* \vdash t' = t''$, then $\Gamma^* \vdash t = t''$.
4. If $\Gamma^* \vdash t = t'$, then

$$\Gamma^* \vdash f(t_1, \dots, t_{i-1}, t, t_{i+1}, \dots, t_n) = f(t_1, \dots, t_{i-1}, t', t_{i+1}, \dots, t_n)$$

for every n -place function symbol f and terms $t_1, \dots, t_{i-1}, t_{i+1}, \dots, t_n$.

5. If $\Gamma^* \vdash t = t'$ and $\Gamma^* \vdash R(t_1, \dots, t_{i-1}, t, t_{i+1}, \dots, t_n)$, then $\Gamma^* \vdash R(t_1, \dots, t_{i-1}, t', t_{i+1}, \dots, t_n)$ for every n -place **predicate symbol** R and terms $t_1, \dots, t_{i-1}, t_{i+1}, \dots, t_n$.

□

Problem 7.4. Complete the proof of [Proposition 7.13](#).

Definition 7.14. Suppose Γ^* is a consistent and **complete** set in a language \mathcal{L} , t is a term, and \approx as in the previous definition. Then:

$$[t]_{\approx} = \{t' : t' \in \text{Trm}(\mathcal{L}), t \approx t'\}$$

and $\text{Trm}(\mathcal{L})/\approx = \{[t]_{\approx} : t \in \text{Trm}(\mathcal{L})\}$.

Definition 7.15. Let $\mathfrak{M} = \mathfrak{M}(\Gamma^*)$ be the term model for Γ^* . Then \mathfrak{M}/\approx is the following **structure**:

fol:com:ide:
defn:term-model-factor

1. $|\mathfrak{M}/\approx| = \text{Trm}(\mathcal{L})/\approx$.
2. $c^{\mathfrak{M}/\approx} = [c]_{\approx}$
3. $f^{\mathfrak{M}/\approx}([t_1]_{\approx}, \dots, [t_n]_{\approx}) = [f(t_1, \dots, t_n)]_{\approx}$
4. $\langle [t_1]_{\approx}, \dots, [t_n]_{\approx} \rangle \in R^{\mathfrak{M}/\approx}$ iff $\mathfrak{M} \models R(t_1, \dots, t_n)$.

explanation

Note that we have defined $f^{\mathfrak{M}/\approx}$ and $R^{\mathfrak{M}/\approx}$ for elements of $\text{Trm}(\mathcal{L})/\approx$ by referring to them as $[t]_{\approx}$, i.e., via *representatives* $t \in [t]_{\approx}$. We have to make sure that these definitions do not depend on the choice of these representatives, i.e., that for some other choices t' which determine the same equivalence classes ($[t]_{\approx} = [t']_{\approx}$), the definitions yield the same result. For instance, if R is a one-place **predicate symbol**, the last clause of the definition says that $[t]_{\approx} \in R^{\mathfrak{M}/\approx}$ iff $\mathfrak{M} \models R(t)$. If for some other term t' with $t \approx t'$, $\mathfrak{M} \not\models R(t)$, then the definition would require $[t']_{\approx} \notin R^{\mathfrak{M}/\approx}$. If $t \approx t'$, then $[t]_{\approx} = [t']_{\approx}$, but we can't have both $[t]_{\approx} \in R^{\mathfrak{M}/\approx}$ and $[t]_{\approx} \notin R^{\mathfrak{M}/\approx}$. However, [Proposition 7.13](#) guarantees that this cannot happen.

Proposition 7.16. \mathfrak{M}/\approx is well defined, i.e., if $t_1, \dots, t_n, t'_1, \dots, t'_n$ are terms, and $t_i \approx t'_i$ then

1. $[f(t_1, \dots, t_n)]_{\approx} = [f(t'_1, \dots, t'_n)]_{\approx}$, i.e.,

$$f(t_1, \dots, t_n) \approx f(t'_1, \dots, t'_n)$$

and

2. $\mathfrak{M} \models R(t_1, \dots, t_n)$ iff $\mathfrak{M} \models R(t'_1, \dots, t'_n)$, i.e.,

$$R(t_1, \dots, t_n) \in \Gamma^* \text{ iff } R(t'_1, \dots, t'_n) \in \Gamma^*.$$

Proof. Follows from [Proposition 7.13](#) by induction on n . □

fol.com:ide: **Lemma 7.17.** $\mathfrak{M}/\approx \models \varphi$ iff $\varphi \in \Gamma^*$ for all sentences φ .
lem:truth

Proof. By induction on φ , just as in the proof of [Lemma 7.11](#). The only case that needs additional attention is when $\varphi \equiv t = t'$.

$$\begin{aligned} \mathfrak{M}/\approx \models t = t' &\text{ iff } [t]_{\approx} = [t']_{\approx} \text{ (by definition of } \mathfrak{M}/\approx) \\ &\text{ iff } t \approx t' \text{ (by definition of } [t]_{\approx}) \\ &\text{ iff } t = t' \in \Gamma^* \text{ (by definition of } \approx). \end{aligned}$$

□

Note that while $\mathfrak{M}(\Gamma^*)$ is always **enumerable** and infinite, \mathfrak{M}/\approx may be **finite**, since it may turn out that there are only finitely many classes $[t]_{\approx}$. This is to be expected, since Γ may contain **sentences** which require any **structure** in which they are true to be finite. For instance, $\forall x \forall y x = y$ is a consistent **sentence**, but is satisfied only in **structures** with a **domain** that contains exactly one **element**. *digression*

7.8 The Completeness Theorem

fol.com:cth: Let's combine our results: we arrive at Gödel's completeness theorem. *explanation*
sec

fol.com:cth: **Theorem 7.18** (Completeness Theorem). *Let Γ be a set of sentences. If Γ is consistent, it is satisfiable.*
thm:completeness

Proof. Suppose Γ is consistent. By [Lemma 7.6](#), there is a saturated consistent set $\Gamma' \supseteq \Gamma$. By [Lemma 7.8](#), there is a $\Gamma^* \supseteq \Gamma'$ which is consistent and **complete**. Since $\Gamma' \subseteq \Gamma^*$, for each **formula** φ , Γ^* contains a **formula** of the form $\exists x \varphi \rightarrow \varphi(c)$ and so Γ^* is saturated.

If Γ does not contain $=$, then by [Lemma 7.11](#), $\mathfrak{M}(\Gamma^*) \models \varphi$ iff $\varphi \in \Gamma^*$. From this it follows in particular that for all $\varphi \in \Gamma$, $\mathfrak{M}(\Gamma^*) \models \varphi$, so Γ is satisfiable. If Γ does contain $=$, then by [Lemma 7.17](#), $\mathfrak{M}/\approx \models \varphi$ iff $\varphi \in \Gamma^*$ for all sentences φ . In particular, $\mathfrak{M}/\approx \models \varphi$ for all $\varphi \in \Gamma$, so Γ is satisfiable. □

fol.com:cth: **Corollary 7.19** (Completeness Theorem, Second Version). *For all Γ and φ sentences: if $\Gamma \models \varphi$ then $\Gamma \vdash \varphi$.*
cor:completeness

Proof. Note that the Γ 's in [Corollary 7.19](#) and [Theorem 7.18](#) are universally quantified. To make sure we do not confuse ourselves, let us restate [Theorem 7.18](#) using a different variable: for any set of sentences Δ , if Δ is consistent, it is satisfiable. By contraposition, if Δ is not satisfiable, then Δ is inconsistent. We will use this to prove the corollary.

Suppose that $\Gamma \models \varphi$. Then $\Gamma \cup \{\neg\varphi\}$ is unsatisfiable by [Proposition 1.50](#). Taking $\Gamma \cup \{\neg\varphi\}$ as our Δ , the previous version of [Theorem 7.18](#) gives us that $\Gamma \cup \{\neg\varphi\}$ is inconsistent. By [Propositions 4.19](#), [5.19](#) and [6.27](#), $\Gamma \vdash \varphi$. □

Problem 7.5. Use [Corollary 7.19](#) to prove [Theorem 7.18](#), thus showing that the two formulations of the completeness theorem are equivalent.

Problem 7.6. In order for a **derivation** system to be complete, its rules must be strong enough to prove every unsatisfiable set inconsistent. Which of the rules of **derivation** were necessary to prove completeness? Are any of these rules not used anywhere in the proof? In order to answer these questions, make a list or diagram that shows which of the rules of **derivation** were used in which results that lead up to the proof of [Theorem 7.18](#). Be sure to note any tacit uses of rules in these proofs.

7.9 The Compactness Theorem

One important consequence of the completeness theorem is the compactness theorem. The compactness theorem states that if each *finite* subset of a set of **sentences** is satisfiable, the entire set is satisfiable—even if the set itself is infinite. This is far from obvious. There is nothing that seems to rule out, at first glance at least, the possibility of there being infinite sets of **sentences** which are contradictory, but the contradiction only arises, so to speak, from the infinite number. The compactness theorem says that such a scenario can be ruled out: there are no unsatisfiable infinite sets of **sentences** each finite subset of which is satisfiable. Like the completeness theorem, it has a version related to entailment: if an infinite set of **sentences** entails something, already a finite subset does.

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Definition 7.20. A set Γ of **formulas** is *finitely satisfiable* if and only if every finite $\Gamma_0 \subseteq \Gamma$ is satisfiable.

Theorem 7.21 (Compactness Theorem). *The following hold for any sentences Γ and φ :*

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thm:compactness

1. $\Gamma \models \varphi$ iff there is a finite $\Gamma_0 \subseteq \Gamma$ such that $\Gamma_0 \models \varphi$.
2. Γ is satisfiable if and only if it is finitely satisfiable.

Proof. We prove (2). If Γ is satisfiable, then there is a **structure** \mathfrak{M} such that $\mathfrak{M} \models \varphi$ for all $\varphi \in \Gamma$. Of course, this \mathfrak{M} also satisfies every finite subset of Γ , so Γ is finitely satisfiable.

Now suppose that Γ is finitely satisfiable. Then every finite subset $\Gamma_0 \subseteq \Gamma$ is satisfiable. By soundness ([Corollaries 5.29, 4.31](#) and [6.38](#)), every finite subset is consistent. Then Γ itself must be consistent by [Propositions 4.17, 5.17](#) and [6.21](#). By completeness ([Theorem 7.18](#)), since Γ is consistent, it is satisfiable. \square

Problem 7.7. Prove (1) of [Theorem 7.21](#).

Example 7.22. In every model \mathfrak{M} of a theory Γ , each term t of course picks out an **element** of $|\mathfrak{M}|$. Can we guarantee that it is also true that every **element** of $|\mathfrak{M}|$ is picked out by some term or other? In other words, are there theories Γ all models of which are covered? The compactness theorem shows that this is not the case if Γ has infinite models. Here's how to see this: Let \mathfrak{M} be an

infinite model of Γ , and let c be a **constant symbol** not in the language of Γ . Let Δ be the set of all sentences $c \neq t$ for t a term in the language \mathcal{L} of Γ , i.e.,

$$\Delta = \{c \neq t : t \in \text{Trm}(\mathcal{L})\}.$$

A finite subset of $\Gamma \cup \Delta$ can be written as $\Gamma' \cup \Delta'$, with $\Gamma' \subseteq \Gamma$ and $\Delta' \subseteq \Delta$. Since Δ' is finite, it can contain only finitely many terms. Let $a \in |\mathfrak{M}|$ be an **element** of $|\mathfrak{M}|$ not picked out by any of them, and let \mathfrak{M}' be the **structure** that is just like \mathfrak{M} , but also $c^{\mathfrak{M}'} = a$. Since $a \neq \text{Val}^{\mathfrak{M}}(t)$ for all t occurring in Δ' , $\mathfrak{M}' \models \Delta'$. Since $\mathfrak{M} \models \Gamma$, $\Gamma' \subseteq \Gamma$, and c does not occur in Γ , also $\mathfrak{M}' \models \Gamma'$. Together, $\mathfrak{M}' \models \Gamma' \cup \Delta'$ for every finite subset $\Gamma' \cup \Delta'$ of $\Gamma \cup \Delta$. So every finite subset of $\Gamma \cup \Delta$ is satisfiable. By compactness, $\Gamma \cup \Delta$ itself is satisfiable. So there are models $\mathfrak{M} \models \Gamma \cup \Delta$. Every such \mathfrak{M} is a model of Γ , but is not covered, since $\text{Val}^{\mathfrak{M}}(c) \neq \text{Val}^{\mathfrak{M}}(t)$ for all terms t of \mathcal{L} .

Example 7.23. Consider a language \mathcal{L} containing the **predicate symbol** $<$, **constant symbols** $0, 1$, and **function symbols** $+, \times, -, \div$. Let Γ be the set of all **sentences** in this language true in \mathfrak{Q} with domain \mathfrak{Q} and the obvious interpretations. Γ is the set of all **sentences** of \mathcal{L} true about the rational numbers. Of course, in \mathfrak{Q} (and even in \mathfrak{R}), there are no numbers which are greater than 0 but less than $1/k$ for all $k \in \mathbb{Z}^+$. Such a number, if it existed, would be an *infinitesimal*: non-zero, but infinitely small. The compactness theorem shows that there are models of Γ in which infinitesimals exist: Let Δ be $\{0 < c\} \cup \{c < (1 \div \bar{k}) : k \in \mathbb{Z}^+\}$ (where $\bar{k} = (1 + (1 + \dots + (1 + 1) \dots))$ with k 1's). For any finite subset Δ_0 of Δ there is a K such that all the **sentences** $c < \bar{k}$ in Δ_0 have $k < K$. If we expand \mathfrak{Q} to \mathfrak{Q}' with $c^{\mathfrak{Q}'} = 1/K$ we have that $\mathfrak{Q}' \models \Gamma \cup \Delta_0$, and so $\Gamma \cup \Delta$ is finitely satisfiable (Exercise: prove this in detail). By compactness, $\Gamma \cup \Delta$ is satisfiable. Any model \mathfrak{S} of $\Gamma \cup \Delta$ contains an infinitesimal, namely $c^{\mathfrak{S}}$.

Problem 7.8. In the standard model of arithmetic \mathfrak{N} , there is no **element** $k \in |\mathfrak{N}|$ which satisfies every formula $\bar{n} < x$ (where \bar{n} is $0' \dots'$ with n 1's). Use the compactness theorem to show that the set of sentences in the language of arithmetic which are true in the standard model of arithmetic \mathfrak{N} are also true in a **structure** \mathfrak{N}' that contains an **element** which *does* satisfy every formula $\bar{n} < x$.

Example 7.24. We know that first-order logic with **identity predicate** can express that the size of the **domain** must have some minimal size: The sentence $\varphi_{\geq n}$ (which says “there are at least n distinct objects”) is true only in structures where $|\mathfrak{M}|$ has at least n objects. So if we take

$$\Delta = \{\varphi_{\geq n} : n \geq 1\}$$

then any model of Δ must be infinite. Thus, we can guarantee that a theory only has infinite models by adding Δ to it: the models of $\Gamma \cup \Delta$ are all and only the infinite models of Γ .

So first-order logic can express infinitude. The compactness theorem shows that it cannot express finitude, however. For suppose some set of sentences A were satisfied in all and only finite **structures**. Then $\Delta \cup A$ is finitely satisfiable. Why? Suppose $\Delta' \cup A' \subseteq \Delta \cup A$ is finite with $\Delta' \subseteq \Delta$ and $A' \subseteq A$. Let n be the largest number such that $A_{\geq n} \in \Delta'$. A , being satisfied in all finite structures, has a model \mathfrak{M} with finitely many but $\geq n$ **elements**. But then $\mathfrak{M} \models \Delta' \cup A'$. By compactness, $\Delta \cup A$ has an infinite model, contradicting the assumption that A is satisfied only in finite **structures**.

7.10 A Direct Proof of the Compactness Theorem

We can prove the Compactness Theorem directly, without appealing to the Completeness Theorem, using the same ideas as in the proof of the completeness theorem. In the proof of the Completeness Theorem we started with a consistent set Γ of **sentences**, expanded it to a consistent, saturated, and **complete** set Γ^* of **sentences**, and then showed that in the term model $\mathfrak{M}(\Gamma^*)$ constructed from Γ^* , all **sentences** of Γ are true, so Γ is satisfiable.

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We can use the same method to show that a finitely satisfiable set of sentences is satisfiable. We just have to prove the corresponding versions of the results leading to the truth lemma where we replace “consistent” with “finitely satisfiable.”

Proposition 7.25. *Suppose Γ is **complete** and **finitely satisfiable**. Then:*

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prop:fsat-ccs

1. $(\varphi \wedge \psi) \in \Gamma$ iff both $\varphi \in \Gamma$ and $\psi \in \Gamma$.
2. $(\varphi \vee \psi) \in \Gamma$ iff either $\varphi \in \Gamma$ or $\psi \in \Gamma$.
3. $(\varphi \rightarrow \psi) \in \Gamma$ iff either $\varphi \notin \Gamma$ or $\psi \in \Gamma$.

Problem 7.9. Prove [Proposition 7.25](#). Avoid the use of \vdash .

Lemma 7.26. *Every **finitely satisfiable** set Γ can be extended to a **saturated** **finitely satisfiable** set Γ' .*

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lem:fsat-henkin

Problem 7.10. Prove [Lemma 7.26](#). (Hint: The crucial step is to show that if Γ_n is finitely satisfiable, so is $\Gamma_n \cup \{\theta_n\}$, without any appeal to **derivations** or consistency.)

Proposition 7.27. *Suppose Γ is **complete**, **finitely satisfiable**, and **saturated**.*

fol:com:cpd:
prop:fsat-instances

1. $\exists x \varphi(x) \in \Gamma$ iff $\varphi(t) \in \Gamma$ for at least one closed term t .
2. $\forall x \varphi(x) \in \Gamma$ iff $\varphi(t) \in \Gamma$ for all closed terms t .

Problem 7.11. Prove [Proposition 7.27](#).

Lemma 7.28. *Every **finitely satisfiable** set Γ' can be extended to a **complete** and **finitely satisfiable** set Γ^* .*

fol:com:cpd:
lem:fsat-lindenbaum

Problem 7.12. Prove [Lemma 7.28](#). (Hint: the crucial step is to show that if Γ_n is finitely satisfiable, then either $\Gamma_n \cup \{\varphi_n\}$ or $\Gamma_n \cup \{\neg\varphi_n\}$ is finitely satisfiable.)

fol:com:cpd:
thm:compactness-direct **Theorem 7.29** (Compactness). *Γ is satisfiable if and only if it is finitely satisfiable.*

Proof. If Γ is satisfiable, then there is a **structure** \mathfrak{M} such that $\mathfrak{M} \models \varphi$ for all $\varphi \in \Gamma$. Of course, this \mathfrak{M} also satisfies every finite subset of Γ , so Γ is finitely satisfiable.

Now suppose that Γ is finitely satisfiable. By [Lemma 7.26](#), there is a finitely satisfiable, saturated set $\Gamma' \supseteq \Gamma$. By [Lemma 7.28](#), Γ' can be extended to a **complete** and finitely satisfiable set Γ^* , and Γ^* is still saturated. Construct the term model $\mathfrak{M}(\Gamma^*)$ as in [Definition 7.9](#). Note that [Proposition 7.10](#) did not rely on the fact that Γ^* is consistent (or **complete** or saturated, for that matter), but just on the fact that $\mathfrak{M}(\Gamma^*)$ is covered. The proof of the Truth Lemma ([Lemma 7.11](#)) goes through if we replace references to [Proposition 7.2](#) and [Proposition 7.7](#) by references to [Proposition 7.25](#) and [Proposition 7.27](#). \square

Problem 7.13. Write out the complete proof of the Truth Lemma ([Lemma 7.11](#)) in the version required for the proof of [Theorem 7.29](#).

7.11 The Löwenheim-Skolem Theorem

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sec The Löwenheim-Skolem Theorem says that if a theory has an infinite model, then it also has a model that is at most **denumerable**. An immediate consequence of this fact is that first-order logic cannot express that the size of a **structure** is **non-enumerable**: any **sentence** or set of **sentences** satisfied in all **non-enumerable structures** is also satisfied in some **enumerable** structure.

fol:com:dls:
thm:downward-ls **Theorem 7.30.** *If Γ is consistent then it has an **enumerable** model, i.e., it is satisfiable in a **structure** whose domain is either finite or **denumerable**.*

Proof. If Γ is consistent, the **structure** \mathfrak{M} delivered by the proof of the completeness theorem has a domain $|\mathfrak{M}|$ that is no larger than the set of the terms of the language \mathcal{L} . So \mathfrak{M} is at most **denumerable**. \square

fol:com:dls:
noidentity-ls **Theorem 7.31.** *If Γ is consistent set of **sentences** in the language of first-order logic without identity, then it has a **denumerable** model, i.e., it is satisfiable in a **structure** whose domain is infinite and **enumerable**.*

Proof. If Γ is consistent and contains no sentences in which identity appears, then the **structure** \mathfrak{M} delivered by the proof of the completeness theorem has a domain $|\mathfrak{M}|$ identical to the set of terms of the language \mathcal{L}' . So \mathfrak{M} is **denumerable**, since $\text{Trm}(\mathcal{L}')$ is. \square

Example 7.32 (Skolem’s Paradox). Zermelo-Fraenkel set theory **ZFC** is a very powerful framework in which practically all mathematical statements can be expressed, including facts about the sizes of sets. So for instance, **ZFC** can prove that the set \mathbb{R} of real numbers is **non-enumerable**, it can prove Cantor’s Theorem that the power set of any set is larger than the set itself, etc. If **ZFC** is consistent, its models are all infinite, and moreover, they all contain **elements** about which the theory says that they are **non-enumerable**, such as the element that makes true the theorem of **ZFC** that the power set of the natural numbers exists. By the Löwenheim-Skolem Theorem, **ZFC** also has **enumerable** models—models that contain “**non-enumerable**” sets but which themselves are **enumerable**.

Chapter 8

Beyond First-order Logic

8.1 Overview

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First-order logic is not the only system of logic of interest: there are many extensions and variations of first-order logic. A logic typically consists of the formal specification of a language, usually, but not always, a deductive system, and usually, but not always, an intended semantics. But the technical use of the term raises an obvious question: what do logics that are not first-order logic have to do with the word “logic,” used in the intuitive or philosophical sense? All of the systems described below are designed to model reasoning of some form or another; can we say what makes them logical?

No easy answers are forthcoming. The word “logic” is used in different ways and in different contexts, and the notion, like that of “truth,” has been analyzed from numerous philosophical stances. For example, one might take the goal of logical reasoning to be the determination of which statements are necessarily true, true a priori, true independent of the interpretation of the nonlogical terms, true by virtue of their form, or true by linguistic convention; and each of these conceptions requires a good deal of clarification. Even if one restricts one’s attention to the kind of logic used in mathematics, there is little agreement as to its scope. For example, in the *Principia Mathematica*, Russell and Whitehead tried to develop mathematics on the basis of logic, in the *logicist* tradition begun by Frege. Their system of logic was a form of higher-type logic similar to the one described below. In the end they were forced to introduce axioms which, by most standards, do not seem purely logical (notably, the axiom of infinity, and the axiom of reducibility), but one might nonetheless hold that some forms of higher-order reasoning should be accepted as logical. In contrast, Quine, whose ontology does not admit “propositions” as legitimate objects of discourse, argues that second-order and higher-order logic are really manifestations of set theory in sheep’s clothing; in other words, systems involving quantification over predicates are not purely logical.

For now, it is best to leave such philosophical issues for a rainy day, and simply think of the systems below as formal idealizations of various kinds of

reasoning, logical or otherwise.

8.2 Many-Sorted Logic

In first-order logic, variables and quantifiers range over a single **domain**. But it is often useful to have multiple (disjoint) **domains**: for example, you might want to have a **domain** of numbers, a **domain** of geometric objects, a **domain** of functions from numbers to numbers, a **domain** of abelian groups, and so on. fol:byd:msl:
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Many-sorted logic provides this kind of framework. One starts with a list of “sorts”—the “sort” of an object indicates the “**domain**” it is supposed to inhabit. One then has **variables** and quantifiers for each sort, and (usually) an **identity predicate** for each sort. Functions and relations are also “typed” by the sorts of objects they can take as arguments. Otherwise, one keeps the usual rules of first-order logic, with versions of the quantifier-rules repeated for each sort.

For example, to study international relations we might choose a language with two sorts of objects, French citizens and German citizens. We might have a unary relation, “drinks wine,” for objects of the first sort; another unary relation, “eats wurst,” for objects of the second sort; and a binary relation, “forms a multinational married couple,” which takes two arguments, where the first argument is of the first sort and the second argument is of the second sort. If we use variables a, b, c to range over French citizens and x, y, z to range over German citizens, then

$$\forall a \forall x [(MarriedTo(a, x) \rightarrow (DrinksWine(a) \vee \neg EatsWurst(x)))]$$

asserts that if any French person is married to a German, either the French person drinks wine or the German doesn’t eat wurst.

Many-sorted logic can be embedded in first-order logic in a natural way, by lumping all the objects of the many-sorted **domains** together into one first-order **domain**, using unary **predicate symbols** to keep track of the sorts, and relativizing quantifiers. For example, the first-order language corresponding to the example above would have unary **predicate symbols** “*German*” and “*French*,” in addition to the other relations described, with the sort requirements erased. A sorted quantifier $\forall x \varphi$, where x is a **variable** of the German sort, translates to

$$\forall x (German(x) \rightarrow \varphi).$$

We need to add axioms that insure that the sorts are separate—e.g., $\forall x \neg(German(x) \wedge French(x))$ —as well as axioms that guarantee that “drinks wine” only holds of objects satisfying the predicate $French(x)$, etc. With these conventions and axioms, it is not difficult to show that many-sorted **sentences** translate to first-order **sentences**, and many-sorted **derivations** translate to first-order **derivations**. Also, many-sorted **structures** “translate” to corresponding first-order **structures** and vice-versa, so we also have a completeness theorem for many-sorted logic.

8.3 Second-Order logic

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The language of second-order logic allows one to quantify not just over a **domain** of individuals, but over relations on that **domain** as well. Given a first-order language \mathcal{L} , for each k one adds **variables** R which range over k -ary relations, and allows quantification over those variables. If R is a **variable** for a k -ary relation, and t_1, \dots, t_k are ordinary (first-order) terms, $R(t_1, \dots, t_k)$ is an atomic **formula**. Otherwise, the set of **formulas** is defined just as in the case of first-order logic, with additional clauses for second-order quantification. Note that we only have the **identity predicate** for first-order terms: if R and S are relation **variables** of the same arity k , we can define $R = S$ to be an abbreviation for

$$\forall x_1 \dots \forall x_k (R(x_1, \dots, x_k) \leftrightarrow S(x_1, \dots, x_k)).$$

The rules for second-order logic simply extend the quantifier rules to the new second order variables. Here, however, one has to be a little bit careful to explain how these variables interact with the **predicate symbols** of \mathcal{L} , and with **formulas** of \mathcal{L} more generally. At the bare minimum, relation variables count as terms, so one has inferences of the form

$$\varphi(R) \vdash \exists R \varphi(R)$$

But if \mathcal{L} is the language of arithmetic with a constant relation symbol $<$, one would also expect the following inference to be valid:

$$x < y \vdash \exists R R(x, y)$$

or for a given **formula** φ ,

$$\varphi(x_1, \dots, x_k) \vdash \exists R R(x_1, \dots, x_k)$$

More generally, we might want to allow inferences of the form

$$\varphi[\lambda \vec{x}. \psi(\vec{x})/R] \vdash \exists R \varphi$$

where $\varphi[\lambda \vec{x}. \psi(\vec{x})/R]$ denotes the result of replacing every atomic **formula** of the form Rt_1, \dots, t_k in φ by $\psi(t_1, \dots, t_k)$. This last rule is equivalent to having a *comprehension schema*, i.e., an axiom of the form

$$\exists R \forall x_1, \dots, x_k (\varphi(x_1, \dots, x_k) \leftrightarrow R(x_1, \dots, x_k)),$$

one for each **formula** φ in the second-order language, in which R is not a free variable. (Exercise: show that if R is allowed to occur in φ , this schema is inconsistent!)

When logicians refer to the “axioms of second-order logic” they usually mean the minimal extension of first-order logic by second-order quantifier rules together with the comprehension schema. But it is often interesting to study weaker subsystems of these axioms and rules. For example, note that in its full

generality the axiom schema of comprehension is *impredicative*: it allows one to assert the existence of a relation $R(x_1, \dots, x_k)$ that is “defined” by a **formula** with second-order quantifiers; and these quantifiers range over the set of all such relations—a set which includes R itself! Around the turn of the twentieth century, a common reaction to Russell’s paradox was to lay the blame on such definitions, and to avoid them in developing the foundations of mathematics. If one prohibits the use of second-order quantifiers in the **formula** φ , one has a *predicative* form of comprehension, which is somewhat weaker.

From the semantic point of view, one can think of a second-order **structure** as consisting of a first-order **structure** for the language, coupled with a set of relations on the **domain** over which the second-order quantifiers range (more precisely, for each k there is a set of relations of arity k). Of course, if comprehension is included in the proof system, then we have the added requirement that there are enough relations in the “second-order part” to satisfy the comprehension axioms—otherwise the proof system is not sound! One easy way to insure that there are enough relations around is to take the second-order part to consist of *all* the relations on the first-order part. Such a **structure** is called *full*, and, in a sense, is really the “intended **structure**” for the language. If we restrict our attention to full **structures** we have what is known as the *full* second-order semantics. In that case, specifying a **structure** boils down to specifying the first-order part, since the contents of the second-order part follow from that implicitly.

To summarize, there is some ambiguity when talking about second-order logic. In terms of the proof system, one might have in mind either

1. A “minimal” second-order proof system, together with some comprehension axioms.
2. The “standard” second-order proof system, with full comprehension.

In terms of the semantics, one might be interested in either

1. The “weak” semantics, where a **structure** consists of a first-order part, together with a second-order part big enough to satisfy the comprehension axioms.
2. The “standard” second-order semantics, in which one considers full **structures** only.

When logicians do not specify the proof system or the semantics they have in mind, they are usually referring to the second item on each list. The advantage to using this semantics is that, as we will see, it gives us categorical descriptions of many natural mathematical structures; at the same time, the proof system is quite strong, and sound for this semantics. The drawback is that the proof system is *not* complete for the semantics; in fact, *no* effectively given proof system is complete for the full second-order semantics. On the other hand, we will see that the proof system *is* complete for the weakened semantics; this

implies that if a sentence is not provable, then there is *some structure*, not necessarily the full one, in which it is false.

The language of second-order logic is quite rich. One can identify unary relations with subsets of the *domain*, and so in particular you can quantify over these sets; for example, one can express induction for the natural numbers with a single axiom

$$\forall R ((R(0) \wedge \forall x (R(x) \rightarrow R(x'))) \rightarrow \forall x R(x)).$$

If one takes the language of arithmetic to have symbols $0, ', +, \times$ and $<$, one can add the following axioms to describe their behavior:

1. $\forall x \neg x' = 0$
2. $\forall x \forall y (s(x) = s(y) \rightarrow x = y)$
3. $\forall x (x + 0) = x$
4. $\forall x \forall y (x + y') = (x + y)'$
5. $\forall x (x \times 0) = 0$
6. $\forall x \forall y (x \times y') = ((x \times y) + x)$
7. $\forall x \forall y (x < y \leftrightarrow \exists z y = (x + z'))$

It is not difficult to show that these axioms, together with the axiom of induction above, provide a categorical description of the *structure* \mathfrak{N} , the standard model of arithmetic, provided we are using the full second-order semantics. Given any *structure* \mathfrak{A} in which these axioms are true, define a function f from \mathbb{N} to the *domain* of \mathfrak{A} using ordinary recursion on \mathbb{N} , so that $f(0) = 0^{\mathfrak{A}}$ and $f(x+1) = \iota^{\mathfrak{A}}(f(x))$. Using ordinary induction on \mathbb{N} and the fact that axioms (1) and (2) hold in \mathfrak{A} , we see that f is *injective*. To see that f is *surjective*, let P be the set of elements of $|\mathfrak{A}|$ that are in the range of f . Since \mathfrak{A} is full, P is in the second-order *domain*. By the construction of f , we know that $0^{\mathfrak{A}}$ is in P , and that P is closed under $\iota^{\mathfrak{A}}$. The fact that the induction axiom holds in \mathfrak{A} (in particular, for P) guarantees that P is equal to the entire first-order *domain* of \mathfrak{A} . This shows that f is a *bijection*. Showing that f is a homomorphism is no more difficult, using ordinary induction on \mathbb{N} repeatedly.

In set-theoretic terms, a function is just a special kind of relation; for example, a unary function f can be identified with a binary relation R satisfying $\forall x \exists y R(x, y)$. As a result, one can quantify over functions too. Using the full semantics, one can then define the class of infinite *structures* to be the class of *structures* \mathfrak{A} for which there is an injective function from the *domain* of \mathfrak{A} to a proper subset of itself:

$$\exists f (\forall x \forall y (f(x) = f(y) \rightarrow x = y) \wedge \exists y \forall x f(x) \neq y).$$

The negation of this sentence then defines the class of finite *structures*.

In addition, one can define the class of well-orderings, by adding the following to the definition of a linear ordering:

$$\forall P (\exists x P(x) \rightarrow \exists x (P(x) \wedge \forall y (y < x \rightarrow \neg P(y)))).$$

This asserts that every non-empty set has a least element, modulo the identification of “set” with “one-place relation”. For another example, one can express the notion of connectedness for graphs, by saying that there is no nontrivial separation of the vertices into disconnected parts:

$$\neg \exists A (\exists x A(x) \wedge \exists y \neg A(y) \wedge \forall w \forall z ((A(w) \wedge \neg A(z)) \rightarrow \neg R(w, z))).$$

For yet another example, you might try as an exercise to define the class of finite **structures** whose **domain** has even size. More strikingly, one can provide a categorical description of the real numbers as a complete ordered field containing the rationals.

In short, second-order logic is much more expressive than first-order logic. That’s the good news; now for the bad. We have already mentioned that there is no effective proof system that is complete for the full second-order semantics. For better or for worse, many of the properties of first-order logic are absent, including compactness and the Löwenheim-Skolem theorems.

On the other hand, if one is willing to give up the full second-order semantics in terms of the weaker one, then the minimal second-order proof system is complete for this semantics. In other words, if we read \vdash as “proves in the minimal system” and \models as “logically implies in the weaker semantics”, we can show that whenever $\Gamma \models \varphi$ then $\Gamma \vdash \varphi$. If one wants to include specific comprehension axioms in the proof system, one has to restrict the semantics to second-order structures that satisfy these axioms: for example, if Δ consists of a set of comprehension axioms (possibly all of them), we have that if $\Gamma \cup \Delta \models \varphi$, then $\Gamma \cup \Delta \vdash \varphi$. In particular, if φ is not provable using the comprehension axioms we are considering, then there is a model of $\neg\varphi$ in which these comprehension axioms nonetheless hold.

The easiest way to see that the completeness theorem holds for the weaker semantics is to think of second-order logic as a many-sorted logic, as follows. One sort is interpreted as the ordinary “first-order” **domain**, and then for each k we have a **domain** of “relations of arity k .” We take the language to have built-in relation symbols “ $true_k(R, x_1, \dots, x_k)$ ” which is meant to assert that R holds of x_1, \dots, x_k , where R is a variable of the sort “ k -ary relation” and x_1, \dots, x_k are objects of the first-order sort.

With this identification, the weak second-order semantics is essentially the usual semantics for many-sorted logic; and we have already observed that many-sorted logic can be embedded in first-order logic. Modulo the translations back and forth, then, the weaker conception of second-order logic is really a form of first-order logic in disguise, where the **domain** contains both “objects” and “relations” governed by the appropriate axioms.

8.4 Higher-Order logic

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Passing from first-order logic to second-order logic enabled us to talk about sets of objects in the first-order **domain**, within the formal language. Why stop there? For example, third-order logic should enable us to deal with sets of sets of objects, or perhaps even sets which contain both objects and sets of objects. And fourth-order logic will let us talk about sets of objects of that kind. As you may have guessed, one can iterate this idea arbitrarily.

In practice, higher-order logic is often **formulated** in terms of functions instead of relations. (Modulo the natural identifications, this difference is inessential.) Given some basic “sorts” A, B, C, \dots (which we will now call “types”), we can create new ones by stipulating

If σ and τ are finite types then so is $\sigma \rightarrow \tau$.

Think of types as syntactic “labels,” which classify the objects we want in our **domain**; $\sigma \rightarrow \tau$ describes those objects that are functions which take objects of type σ to objects of type τ . For example, we might want to have a type Ω of truth values, “true” and “false,” and a type \mathbb{N} of natural numbers. In that case, you can think of objects of type $\mathbb{N} \rightarrow \Omega$ as unary relations, or subsets of \mathbb{N} ; objects of type $\mathbb{N} \rightarrow \mathbb{N}$ are functions from natural numbers to natural numbers; and objects of type $(\mathbb{N} \rightarrow \mathbb{N}) \rightarrow \mathbb{N}$ are “functionals,” that is, higher-type functions that take functions to numbers.

As in the case of second-order logic, one can think of higher-order logic as a kind of many-sorted logic, where there is a sort for each type of object we want to consider. But it is usually clearer just to define the syntax of higher-type logic from the ground up. For example, we can define a set of finite types inductively, as follows:

1. \mathbb{N} is a finite type.
2. If σ and τ are finite types, then so is $\sigma \rightarrow \tau$.
3. If σ and τ are finite types, so is $\sigma \times \tau$.

Intuitively, \mathbb{N} denotes the type of the natural numbers, $\sigma \rightarrow \tau$ denotes the type of functions from σ to τ , and $\sigma \times \tau$ denotes the type of pairs of objects, one from σ and one from τ . We can then define a set of terms inductively, as follows:

1. For each type σ , there is a stock of variables x, y, z, \dots of type σ
2. o is a term of type \mathbb{N}
3. S (successor) is a term of type $\mathbb{N} \rightarrow \mathbb{N}$
4. If s is a term of type σ , and t is a term of type $\mathbb{N} \rightarrow (\sigma \rightarrow \sigma)$, then R_{st} is a term of type $\mathbb{N} \rightarrow \sigma$

5. If s is a term of type $\tau \rightarrow \sigma$ and t is a term of type τ , then $s(t)$ is a term of type σ
6. If s is a term of type σ and x is a variable of type τ , then $\lambda x. s$ is a term of type $\tau \rightarrow \sigma$.
7. If s is a term of type σ and t is a term of type τ , then $\langle s, t \rangle$ is a term of type $\sigma \times \tau$.
8. If s is a term of type $\sigma \times \tau$ then $p_1(s)$ is a term of type σ and $p_2(s)$ is a term of type τ .

Intuitively, R_{st} denotes the function defined recursively by

$$\begin{aligned} R_{st}(0) &= s \\ R_{st}(x+1) &= t(x, R_{st}(x)), \end{aligned}$$

$\langle s, t \rangle$ denotes the pair whose first component is s and whose second component is t , and $p_1(s)$ and $p_2(s)$ denote the first and second elements (“projections”) of s . Finally, $\lambda x. s$ denotes the function f defined by

$$f(x) = s$$

for any x of type σ ; so item (6) gives us a form of comprehension, enabling us to define functions using terms. **Formulas** are built up from **identity predicate** statements $s = t$ between terms of the same type, the usual propositional connectives, and higher-type quantification. One can then take the axioms of the system to be the basic equations governing the terms defined above, together with the usual rules of logic with quantifiers and **identity predicate**.

If one augments the finite type system with a type Ω of truth values, one has to include axioms which govern its use as well. In fact, if one is clever, one can get rid of complex **formulas** entirely, replacing them with terms of type Ω ! The proof system can then be modified accordingly. The result is essentially the *simple theory of types* set forth by Alonzo Church in the 1930s.

As in the case of second-order logic, there are different versions of higher-type semantics that one might want to use. In the full version, variables of type $\sigma \rightarrow \tau$ range over the set of *all* functions from the objects of type σ to objects of type τ . As you might expect, this semantics is too strong to admit a complete, effective proof system. But one can consider a weaker semantics, in which a **structure** consists of sets of elements T_τ for each type τ , together with appropriate operations for application, projection, etc. If the details are carried out correctly, one can obtain completeness theorems for the kinds of proof systems described above.

Higher-type logic is attractive because it provides a framework in which we can embed a good deal of mathematics in a natural way: starting with \mathbb{N} , one can define real numbers, continuous functions, and so on. It is also particularly attractive in the context of intuitionistic logic, since the types have clear “constructive” interpretations. In fact, one can develop constructive

versions of higher-type semantics (based on intuitionistic, rather than classical logic) that clarify these constructive interpretations quite nicely, and are, in many ways, more interesting than the classical counterparts.

8.5 Intuitionistic Logic

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In contrast to second-order and higher-order logic, intuitionistic first-order logic represents a restriction of the classical version, intended to model a more “constructive” kind of reasoning. The following examples may serve to illustrate some of the underlying motivations.

Suppose someone came up to you one day and announced that they had determined a natural number x , with the property that if x is prime, the Riemann hypothesis is true, and if x is composite, the Riemann hypothesis is false. Great news! Whether the Riemann hypothesis is true or not is one of the big open questions of mathematics, and here they seem to have reduced the problem to one of calculation, that is, to the determination of whether a specific number is prime or not.

What is the magic value of x ? They describe it as follows: x is the natural number that is equal to 7 if the Riemann hypothesis is true, and 9 otherwise.

Angrily, you demand your money back. From a classical point of view, the description above does in fact determine a unique value of x ; but what you really want is a value of x that is given *explicitly*.

To take another, perhaps less contrived example, consider the following question. We know that it is possible to raise an irrational number to a rational power, and get a rational result. For example, $\sqrt{2}^2 = 2$. What is less clear is whether or not it is possible to raise an irrational number to an *irrational* power, and get a rational result. The following theorem answers this in the affirmative:

Theorem 8.1. *There are irrational numbers a and b such that a^b is rational.*

Proof. Consider $\sqrt{2}^{\sqrt{2}}$. If this is rational, we are done: we can let $a = b = \sqrt{2}$. Otherwise, it is irrational. Then we have

$$(\sqrt{2}^{\sqrt{2}})^{\sqrt{2}} = \sqrt{2}^{\sqrt{2} \cdot \sqrt{2}} = \sqrt{2}^2 = 2,$$

which is certainly rational. So, in this case, let a be $\sqrt{2}^{\sqrt{2}}$, and let b be $\sqrt{2}$. \square

Does this constitute a valid proof? Most mathematicians feel that it does. But again, there is something a little bit unsatisfying here: we have proved the existence of a pair of real numbers with a certain property, without being able to say *which* pair of numbers it is. It is possible to prove the same result, but in such a way that the pair a, b is given in the proof: take $a = \sqrt{3}$ and $b = \log_3 4$. Then

$$a^b = \sqrt{3}^{\log_3 4} = 3^{1/2 \cdot \log_3 4} = (3^{\log_3 4})^{1/2} = 4^{1/2} = 2,$$

since $3^{\log_3 x} = x$.

Intuitionistic logic is designed to model a kind of reasoning where moves like the one in the first proof are disallowed. Proving the existence of an x satisfying $\varphi(x)$ means that you have to give a specific x , and a proof that it satisfies φ , like in the second proof. Proving that φ or ψ holds requires that you can prove one or the other.

Formally speaking, intuitionistic first-order logic is what you get if you omit restrict a proof system for first-order logic in a certain way. Similarly, there are intuitionistic versions of second-order or higher-order logic. From the mathematical point of view, these are just formal deductive systems, but, as already noted, they are intended to model a kind of mathematical reasoning. One can take this to be the kind of reasoning that is justified on a certain philosophical view of mathematics (such as Brouwer's intuitionism); one can take it to be a kind of mathematical reasoning which is more "concrete" and satisfying (along the lines of Bishop's constructivism); and one can argue about whether or not the formal description captures the informal motivation. But whatever philosophical positions we may hold, we can study intuitionistic logic as a formally presented logic; and for whatever reasons, many mathematical logicians find it interesting to do so.

There is an informal constructive interpretation of the intuitionist connectives, usually known as the Brouwer-Heyting-Kolmogorov interpretation. It runs as follows: a proof of $\varphi \wedge \psi$ consists of a proof of φ paired with a proof of ψ ; a proof of $\varphi \vee \psi$ consists of either a proof of φ , or a proof of ψ , where we have explicit information as to which is the case; a proof of $\varphi \rightarrow \psi$ consists of a procedure, which transforms a proof of φ to a proof of ψ ; a proof of $\forall x \varphi(x)$ consists of a procedure which returns a proof of $\varphi(x)$ for any value of x ; and a proof of $\exists x \varphi(x)$ consists of a value of x , together with a proof that this value satisfies φ . One can describe the interpretation in computational terms known as the "Curry-Howard isomorphism" or the "formulas-as-types paradigm": think of a **formula** as specifying a certain kind of data type, and proofs as computational objects of these data types that enable us to see that the corresponding **formula** is true.

Intuitionistic logic is often thought of as being classical logic "minus" the law of the excluded middle. This following theorem makes this more precise.

Theorem 8.2. *Intuitionistically, the following axiom schemata are equivalent:*

1. $(\varphi \rightarrow \perp) \rightarrow \neg\varphi$.
2. $\varphi \vee \neg\varphi$
3. $\neg\neg\varphi \rightarrow \varphi$

Obtaining instances of one schema from either of the others is a good exercise in intuitionistic logic.

The first deductive systems for intuitionistic propositional logic, put forth as formalizations of Brouwer's intuitionism, are due, independently, to Kol-

mogorov, Glivenko, and Heyting. The first formalization of intuitionistic first-order logic (and parts of intuitionist mathematics) is due to Heyting. Though a number of classically valid schemata are not intuitionistically valid, many are.

The *double-negation translation* describes an important relationship between classical and intuitionist logic. It is defined inductively as follows (think of φ^N as the “intuitionist” translation of the classical formula φ):

$$\begin{aligned}\varphi^N &\equiv \neg\neg\varphi \quad \text{for atomic formula } \varphi \\ (\varphi \wedge \psi)^N &\equiv (\varphi^N \wedge \psi^N) \\ (\varphi \vee \psi)^N &\equiv \neg\neg(\varphi^N \vee \psi^N) \\ (\varphi \rightarrow \psi)^N &\equiv (\varphi^N \rightarrow \psi^N) \\ (\forall x \varphi)^N &\equiv \forall x \varphi^N \\ (\exists x \varphi)^N &\equiv \neg\neg\exists x \varphi^N\end{aligned}$$

Kolmogorov and Glivenko had versions of this translation for propositional logic; for predicate logic, it is due to Gödel and Gentzen, independently. We have

Theorem 8.3. 1. $\varphi \leftrightarrow \varphi^N$ is provable classically

2. If φ is provable classically, then φ^N is provable intuitionistically.

We can now envision the following dialogue. Classical mathematician: “I’ve proved φ !” Intuitionist mathematician: “Your proof isn’t valid. What you’ve really proved is φ^N .” Classical mathematician: “Fine by me!” As far as the classical mathematician is concerned, the intuitionist is just splitting hairs, since the two are equivalent. But the intuitionist insists there is a difference.

Note that the above translation concerns pure logic only; it does not address the question as to what the appropriate *nonlogical* axioms are for classical and intuitionistic mathematics, or what the relationship is between them. But the following slight extension of the theorem above provides some useful information:

Theorem 8.4. If Γ proves φ classically, Γ^N proves φ^N intuitionistically.

In other words, if φ is provable from some hypotheses classically, then φ^N is provable from their double-negation translations.

To show that a sentence or propositional formula is intuitionistically valid, all you have to do is provide a proof. But how can you show that it is not valid? For that purpose, we need a semantics that is sound, and preferably complete. A semantics due to Kripke nicely fits the bill.

We can play the same game we did for classical logic: define the semantics, and prove soundness and completeness. It is worthwhile, however, to note the following distinction. In the case of classical logic, the semantics was the “obvious” one, in a sense implicit in the meaning of the connectives. Though

one can provide some intuitive motivation for Kripke semantics, the latter does not offer the same feeling of inevitability. In addition, the notion of a classical **structure** is a natural mathematical one, so we can either take the notion of a **structure** to be a tool for studying classical first-order logic, or take classical first-order logic to be a tool for studying mathematical **structures**. In contrast, Kripke **structures** can only be viewed as a logical construct; they don't seem to have independent mathematical interest.

A Kripke **structure** for a propositional language consists of a partial order $\text{Mod}(P)$ with a least element, and an “monotone” assignment of propositional variables to the elements of $\text{Mod}(P)$. The intuition is that the elements of $\text{Mod}(P)$ represent “worlds,” or “states of knowledge”; an element $p \geq q$ represents a “possible future state” of q ; and the propositional variables assigned to p are the propositions that are known to be true in state p . The forcing relation $\mathfrak{F}, p \models \varphi$ then extends this relationship to arbitrary **formulas** in the language; read $\mathfrak{F}, p \models \varphi$ as “ φ is true in state p .” The relationship is defined inductively, as follows:

1. $\mathfrak{F}, p \models p_i$ iff p_i is one of the propositional variables assigned to p .
2. $\mathfrak{F}, p \not\models \perp$.
3. $\mathfrak{F}, p \models (\varphi \wedge \psi)$ iff $\mathfrak{F}, p \models \varphi$ and $\mathfrak{F}, p \models \psi$.
4. $\mathfrak{F}, p \models (\varphi \vee \psi)$ iff $\mathfrak{F}, p \models \varphi$ or $\mathfrak{F}, p \models \psi$.
5. $\mathfrak{F}, p \models (\varphi \rightarrow \psi)$ iff, whenever $q \geq p$ and $\mathfrak{F}, q \models \varphi$, then $\mathfrak{F}, q \models \psi$.

It is a good exercise to try to show that $\neg(p \wedge q) \rightarrow (\neg p \vee \neg q)$ is not intuitionistically valid, by cooking up a Kripke **structure** that provides a counterexample.

8.6 Modal Logics

Consider the following example of a conditional sentence:

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If Jeremy is alone in that room, then he is drunk and naked and dancing on the chairs.

This is an example of a conditional assertion that may be materially true but nonetheless misleading, since it seems to suggest that there is a stronger link between the antecedent and conclusion other than simply that either the antecedent is false or the consequent true. That is, the wording suggests that the claim is not only true in this particular world (where it may be trivially true, because Jeremy is not alone in the room), but that, moreover, the conclusion *would have* been true *had* the antecedent been true. In other words, one can take the assertion to mean that the claim is true not just in this world, but in any “possible” world; or that it is *necessarily* true, as opposed to just true in this particular world.

Modal logic was designed to make sense of this kind of necessity. One obtains modal propositional logic from ordinary propositional logic by adding a box operator; which is to say, if φ is a **formula**, so is $\Box\varphi$. Intuitively, $\Box\varphi$ asserts that φ is *necessarily* true, or true in any possible world. $\Diamond\varphi$ is usually taken to be an abbreviation for $\neg\Box\neg\varphi$, and can be read as asserting that φ is *possibly* true. Of course, modality can be added to predicate logic as well.

Kripke **structures** can be used to provide a semantics for modal logic; in fact, Kripke first designed this semantics with modal logic in mind. Rather than restricting to partial orders, more generally one has a set of “possible worlds,” P , and a binary “accessibility” relation $R(x, y)$ between worlds. Intuitively, $R(p, q)$ asserts that the world q is compatible with p ; i.e., if we are “in” world p , we have to entertain the possibility that the world could have been like q .

Modal logic is sometimes called an “intensional” logic, as opposed to an “extensional” one. The intended semantics for an extensional logic, like classical logic, will only refer to a single world, the “actual” one; while the semantics for an “intensional” logic relies on a more elaborate ontology. In addition to **structuring** necessity, one can use modality to **structure** other linguistic constructions, reinterpreting \Box and \Diamond according to the application. For example:

1. In provability logic, $\Box\varphi$ is read “ φ is provable” and $\Diamond\varphi$ is read “ φ is consistent.”
2. In epistemic logic, one might read $\Box\varphi$ as “I know φ ” or “I believe φ .”
3. In temporal logic, one can read $\Box\varphi$ as “ φ is always true” and $\Diamond\varphi$ as “ φ is sometimes true.”

One would like to augment logic with rules and axioms dealing with modality. For example, the system **S4** consists of the ordinary axioms and rules of propositional logic, together with the following axioms:

$$\begin{aligned} \Box(\varphi \rightarrow \psi) &\rightarrow (\Box\varphi \rightarrow \Box\psi) \\ \Box\varphi &\rightarrow \varphi \\ \Box\varphi &\rightarrow \Box\Box\varphi \end{aligned}$$

as well as a rule, “from φ conclude $\Box\varphi$.” **S5** adds the following axiom:

$$\Diamond\varphi \rightarrow \Box\Diamond\varphi$$

Variations of these axioms may be suitable for different applications; for example, S5 is usually taken to characterize the notion of logical necessity. And the nice thing is that one can usually find a semantics for which the proof system is sound and complete by restricting the accessibility relation in the Kripke **structures** in natural ways. For example, **S4** corresponds to the class of Kripke **structures** in which the accessibility relation is reflexive and transitive. **S5** corresponds to the class of Kripke **structures** in which the accessibility relation is *universal*, which is to say that every world is accessible from every other; so $\Box\varphi$ holds if and only if φ holds in every world.

8.7 Other Logics

As you may have gathered by now, it is not hard to design a new logic. You too can create your own a syntax, make up a deductive system, and fashion a semantics to go with it. You might have to be a bit clever if you want the proof system to be complete for the semantics, and it might take some effort to convince the world at large that your logic is truly interesting. But, in return, you can enjoy hours of good, clean fun, exploring your logic’s mathematical and computational properties.

Recent decades have witnessed a veritable explosion of formal logics. Fuzzy logic is designed to model reasoning about vague properties. Probabilistic logic is designed to model reasoning about uncertainty. Default logics and nonmonotonic logics are designed to model defeasible forms of reasoning, which is to say, “reasonable” inferences that can later be overturned in the face of new information. There are epistemic logics, designed to model reasoning about knowledge; causal logics, designed to model reasoning about causal relationships; and even “deontic” logics, which are designed to model reasoning about moral and ethical obligations. Depending on whether the primary motivation for introducing these systems is philosophical, mathematical, or computational, you may find such creatures studies under the rubric of mathematical logic, philosophical logic, artificial intelligence, cognitive science, or elsewhere.

The list goes on and on, and the possibilities seem endless. We may never attain Leibniz’ dream of reducing all of human reason to calculation—but that can’t stop us from trying.

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Bibliography