

Part I

Normal Modal Logics

This part covers the metatheory of normal modal logics. It currently consists of Aldo Antonelli's notes on classical correspondence theory for basic modal logic.

Chapter 1

Syntax and Semantics of Normal Modal Logics

1.1 Introduction

mod:syn:int:
sec

Modal Logic deals with *modal propositions* and the entailment relations among them. Examples of modal propositions are the following:

1. It is necessary that $2 + 2 = 4$.
2. It is necessarily possible that it will rain tomorrow.
3. If it is necessarily possible that φ then it is possible that φ .

Possibility and necessity are not the only modalities: other unary connectives are also classified as modalities, for instance, “it ought to be the case that φ ,” “It will be the case that φ ,” “Dana knows that φ ,” or “Dana believes that φ .”

Modal logic makes its first appearance in Aristotle’s *De Interpretatione*: he was the first to notice that necessity implies possibility, but not vice versa; that possibility and necessity are inter-definable; that If $\varphi \wedge \psi$ is possibly true then φ is possibly true and ψ is possibly true, but not conversely; and that if $\varphi \rightarrow \psi$ is necessary, then if φ is necessary, so is ψ .

The first modern approach to modal logic was the work of C. I. Lewis, culminating with Lewis and Langford, *Symbolic Logic* (1932). Lewis & Langford were unhappy with the representation of implication by means of the material conditional: $\varphi \rightarrow \psi$ is a poor substitute for “ φ implies ψ .” Instead, they proposed to characterize implication as “Necessarily, if φ then ψ ,” symbolized as $\varphi \rightarrow \psi$. In trying to sort out the different properties, Lewis identified five different modal systems, **S1**, . . . , **S4**, **S5**, the last two of which are still in use.

The approach of Lewis and Langford was purely *syntactical*: they identified reasonable axioms and rules and investigated what was provable with those means. A semantic approach remained elusive for a long time, until a first attempt was made by Rudolf Carnap in *Meaning and Necessity* (1947) using the notion of a *state description*, i.e., a collection of atomic sentences (those

that are “true” in that state description). After lifting the truth definition to arbitrary sentences φ , Carnap defines φ to be *necessarily true* if it is true in all state descriptions. Carnap’s approach could not handle *iterated* modalities, in that sentences of the form “Possibly necessarily . . . possibly φ ” always reduce to the innermost modality.

The major breakthrough in modal semantics came with Saul Kripke’s article “A Completeness Theorem in Modal Logic” (JSL 1959). Kripke based his work on Leibniz’s idea that a statement is necessarily true if it is true “at all possible worlds.” This idea, though, suffers from the same drawbacks as Carnap’s, in that the truth of statement at a world w (or a state description s) does not depend on w at all. So Kripke assumed that worlds are related by an *accessibility relation* R , and that a statement of the form “Necessarily φ ” is true at a world w if and only if φ is true at all worlds w' *accessible from* w . Semantics that provide some version of this approach are called Kripke semantics and made possible the tumultuous development of modal logics (in the plural).

When interpreted by the Kripke semantics, modal logic shows us what *relational structures* look like “from the inside.” A relational structure is just a set equipped with a binary relation (for instance, the set of students in the class ordered by their social security number is a relational structure). But in fact relational structures come in all sorts of domains: besides relative possibility of states of the world, we can have epistemic states of some agent related by epistemic possibility, or states of a dynamical system with their state transitions, etc. Modal logic can be used to model all of these: the first give us ordinary, alethic, modal logic; the others give us epistemic logic, dynamic logic, etc.

We focus on one particular angle, known to modal logicians as “correspondence theory.” One of the most significant early discoveries of Kripke’s is that many properties of the accessibility relation R (whether it is transitive, symmetric, etc.) can be characterized *in the modal language* itself by means of appropriate “modal schemas.” Modal logicians say, for instance, that the reflexivity of R “corresponds” to the schema “If necessarily φ , then φ ”. We explore mainly the correspondence theory of a number of classical systems of modal logic (e.g., **S4** and **S5**) obtained by a combination of the schemas D, T, B, 4, and 5.

1.2 The Language of Basic Modal Logic

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Definition 1.1. The basic language of modal logic contains

1. The propositional constant for **falsity** \perp .
2. The propositional constant for **truth** \top .
3. A **denumerable** set of **propositional variables**: p_0, p_1, p_2, \dots

4. The propositional connectives: \neg (negation), \wedge (conjunction), \vee (disjunction), \rightarrow (conditional), \leftrightarrow (biconditional).
5. The modal operator \Box .
6. The modal operator \Diamond .

Definition 1.2. *Formulas* of the basic modal language are inductively defined as follows:

1. \perp is an atomic formula.
2. \top is an atomic formula.
3. Every propositional variable p_i is an (atomic) formula.
4. If φ and ψ are formulas, then $(\varphi \wedge \psi)$ is a formula.
5. If φ and ψ are formulas, then $(\varphi \vee \psi)$ is a formula.
6. If φ and ψ are formulas, then $(\varphi \rightarrow \psi)$ is a formula.
7. If φ and ψ are formulas, then $(\varphi \leftrightarrow \psi)$ is a formula.
8. If φ is a formula, so is $\Box\varphi$.
9. If φ is a formula, then $\Diamond\varphi$ is a formula.
10. Nothing else is a formula.

If a formula φ does not contain \Box or \Diamond , we say it is *modal-free*.

1.3 Simultaneous Substitution

mod:syn:sub:sec An *instance* of a formula φ is the result of replacing all occurrences of a propositional variable in φ by some other formula. We will refer to instances of formulas often, both when discussing validity and when discussing derivability. It therefore is useful to define the notion precisely.

mod:syn:sub:def:subst-inst **Definition 1.3.** Where φ is a modal formula all of whose propositional variables are among p_1, \dots, p_n , and $\theta_1, \dots, \theta_n$ are also modal formulas, we define $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ as the result of simultaneously substituting each θ_i for p_i in φ . Formally, this is a definition by induction on φ :

1. $\varphi \equiv \perp$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is \perp .
2. $\varphi \equiv \top$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is \top .
3. $\varphi \equiv q$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is q , provided $q \neq p_i$ for $i = 1, \dots, n$.
4. $\varphi \equiv p_i$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is θ_i .

5. $\varphi \equiv \neg\psi$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is $\neg\psi[\theta_1/p_1, \dots, \theta_n/p_n]$.
6. $\varphi \equiv (\psi \wedge \chi)$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is
 $(\psi[\theta_1/p_1, \dots, \theta_n/p_n] \wedge \chi[\theta_1/p_1, \dots, \theta_n/p_n])$.
7. $\varphi \equiv (\psi \vee \chi)$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is
 $(\psi[\theta_1/p_1, \dots, \theta_n/p_n] \vee \chi[\theta_1/p_1, \dots, \theta_n/p_n])$.
8. $\varphi \equiv (\psi \rightarrow \chi)$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is
 $(\psi[\theta_1/p_1, \dots, \theta_n/p_n] \rightarrow \chi[\theta_1/p_1, \dots, \theta_n/p_n])$.
9. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is
 $(\psi[\theta_1/p_1, \dots, \theta_n/p_n] \leftrightarrow \chi[\theta_1/p_1, \dots, \theta_n/p_n])$.
10. $\varphi \equiv \Box\psi$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is $\Box\psi[\theta_1/p_1, \dots, \theta_n/p_n]$.
11. $\varphi \equiv \Diamond\psi$: $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is $\Diamond\psi[\theta_1/p_1, \dots, \theta_n/p_n]$.

The **formula** $\varphi[\theta_1/p_1, \dots, \theta_n/p_n]$ is called a *substitution instance* of φ .

Example 1.4. Suppose φ is $p_1 \rightarrow \Box(p_1 \wedge p_2)$, θ_1 is $\Diamond(p_2 \rightarrow p_3)$ and D_2 is $\neg\Box p_1$. Then $\varphi[\theta_1/p_1, \theta_2/p_2]$ is

$$\Diamond(p_2 \rightarrow p_3) \rightarrow \Box(\Diamond(p_2 \rightarrow p_3) \wedge \neg\Box p_1)$$

while $\varphi[\theta_2/p_1, \theta_1/p_2]$ is

$$\neg\Box p_1 \rightarrow \Box(\neg\Box p_1 \wedge \Diamond(p_2 \rightarrow p_3))$$

Note that simultaneous substitution is in general not the same as iterated substitution, e.g., compare $\varphi[\theta_1/p_1, \theta_2/p_2]$ above with $\varphi[\theta_1/p_1][\theta_2/p_2]$:

$$\Diamond(\neg\Box p_1 \rightarrow p_3) \rightarrow \Box(\Diamond(\neg\Box p_1 \rightarrow p_3) \wedge \neg\Box p_1)$$

and with $\varphi[\theta_2/p_2][\theta_1/p_1]$:

$$\Diamond(\neg\Box p_1 \rightarrow p_3) \rightarrow \Box(\Diamond(\neg\Box p_1 \rightarrow p_3) \wedge \neg\Box\Diamond(\neg\Box p_1 \rightarrow p_3))$$

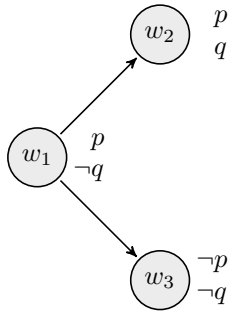


Figure 1.1: A simple model.

mod:syn:rel:
fig:simple

1.4 Relational Models

mod:syn:rel:
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The basic concept of semantics for normal modal logics is that of a *relational model*. It consists of a set of worlds, which are related by a binary “accessibility relation,” together with an assignment which determines which **propositional variables** count as “true” at which worlds.

Definition 1.5. A *model* for the basic modal language is a triple $\mathfrak{M} = \langle W, R, V \rangle$, where

1. W is a nonempty set of “worlds,”
2. R is a binary accessibility relation on W , and
3. V is a function assigning to each **propositional variable** p a set $V(p)$ of possible worlds.

When Rww' holds, we say that w' is *accessible from* w . When $w \in V(p)$ we say p is *true at* w .

The great advantage of relational semantics is that models can be represented by means of simple diagrams, such as the one in [Figure 1.1](#). Worlds are represented by nodes, and world w' is accessible from w precisely when there is an arrow from w to w' . Moreover, we label a node (world) by p when $w \in V(p)$, and otherwise by $\neg p$. [Figure 1.1](#) represents the model with $W = \{w_1, w_2, w_3\}$, $R = \{\langle w_1, w_2 \rangle, \langle w_1, w_3 \rangle\}$, $V(p) = \{w_1, w_2\}$, and $V(q) = \{w_2\}$.

1.5 Truth at a World

mod:syn:trw:
sec

Every modal model determines which modal **formulas** count as true at which worlds in it. The relation “model \mathfrak{M} makes **formula** φ true at world w ” is the basic notion of relational semantics. The relation is defined inductively and coincides with the usual characterization using truth tables for the non-modal operators.

Definition 1.6. *Truth of a formula* φ at w in a \mathfrak{M} , in symbols: $\mathfrak{M}, w \Vdash \varphi$, is defined inductively as follows: mod:syn:trw:
defn:mmodels

1. $\varphi \equiv \perp$: Never $\mathfrak{M}, w \Vdash \perp$.
2. $\varphi \equiv \top$: Always $\mathfrak{M}, w \Vdash \top$.
3. $\mathfrak{M}, w \Vdash p$ iff $w \in V(p)$
4. $\varphi \equiv \neg\psi$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \not\Vdash \psi$.
5. $\varphi \equiv (\psi \wedge \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \Vdash \psi$ and $\mathfrak{M}, w \Vdash \chi$.
6. $\varphi \equiv (\psi \vee \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \Vdash \psi$ or $\mathfrak{M}, w \Vdash \chi$ (or both).
7. $\varphi \equiv (\psi \rightarrow \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \not\Vdash \psi$ or $\mathfrak{M}, w \Vdash \chi$.
8. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff either both $\mathfrak{M}, w \Vdash \psi$ and $\mathfrak{M}, w \Vdash \chi$ or neither $\mathfrak{M}, w \Vdash \psi$ nor $\mathfrak{M}, w \Vdash \chi$.
9. $\varphi \equiv \Box\psi$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w' \Vdash \psi$ for all $w' \in W$ with Rww' mod:syn:trw:
defn:sub:mmodels-box
10. $\varphi \equiv \Diamond\psi$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w' \Vdash \psi$ for at least one $w' \in W$ with Rww' mod:syn:trw:
defn:sub:mmodels-diamond

Note that by clause (9), a formula $\Box\psi$ is true at w whenever there are no w' with wRw' . In such a case $\Box\psi$ is *vacuously* true at w . Also, $\Box\psi$ may be satisfied at w even if ψ is not. The truth of ψ at w does not guarantee the truth of $\Diamond\psi$ at w . This holds, however, if Rww , e.g., if R is reflexive. If there is no w' such that Rww' , then $\mathfrak{M}, w \not\Vdash \Diamond\varphi$, for any φ .

Problem 1.1. Consider the model of Figure 1.1. Which of the following hold?

1. $\mathfrak{M}, w_1 \Vdash q$;
2. $\mathfrak{M}, w_3 \Vdash \neg q$;
3. $\mathfrak{M}, w_1 \Vdash p \vee q$;
4. $\mathfrak{M}, w_1 \Vdash \Box(p \vee q)$;
5. $\mathfrak{M}, w_3 \Vdash \Box q$;
6. $\mathfrak{M}, w_3 \Vdash \Box\perp$;
7. $\mathfrak{M}, w_1 \Vdash \Diamond q$;
8. $\mathfrak{M}, w_1 \Vdash \Box q$;
9. $\mathfrak{M}, w_1 \Vdash \neg\Box\Box\neg q$.

Proposition 1.7.

1. $\mathfrak{M}, w \Vdash \Box\varphi$ iff $\mathfrak{M}, w \Vdash \neg\Diamond\neg\varphi$.

mod:syn:trw:
prop:dual

2. $\mathfrak{M}, w \Vdash \Diamond\varphi$ iff $\mathfrak{M}, w \Vdash \neg\Box\neg\varphi$.

Proof. 1. $\mathfrak{M}, w \Vdash \neg\Box\neg\varphi$ iff $\mathfrak{M} \not\Vdash \Box\neg\varphi$ by definition of $\mathfrak{M}, w \Vdash$. $\mathfrak{M}, w \Vdash \Diamond\neg\varphi$ iff for some w' with Rww' , $\mathfrak{M}, w' \Vdash \neg\varphi$. Hence, $\mathfrak{M}, w \Vdash \Diamond\neg\varphi$ iff for all w' with Rww' , $\mathfrak{M}, w' \not\Vdash \neg\varphi$. We also have $\mathfrak{M}, w' \not\Vdash \neg\varphi$ iff $\mathfrak{M}, w' \Vdash \varphi$. Together we have $\mathfrak{M}, w \Vdash \neg\Box\neg\varphi$ iff for all w' with Rww' , $\mathfrak{M}, w' \Vdash \varphi$. Again by definition of $\mathfrak{M}, w \Vdash$, that is the case iff $\mathfrak{M}, w \Vdash \Box\varphi$.

2. $\mathfrak{M}, w \Vdash \neg\Box\neg\varphi$ iff $\mathfrak{M} \not\Vdash \Box\neg\varphi$. $\mathfrak{M}, w \Vdash \Box\neg\varphi$ iff for all w' with Rww' , $\mathfrak{M}, w' \Vdash \neg\varphi$. Hence, $\mathfrak{M}, w \Vdash \Box\neg\varphi$ iff for some w' with Rww' , $\mathfrak{M}, w' \not\Vdash \neg\varphi$. We also have $\mathfrak{M}, w' \not\Vdash \neg\varphi$ iff $\mathfrak{M}, w' \Vdash \varphi$. Together we have $\mathfrak{M}, w \Vdash \neg\Box\neg\varphi$ iff for some w' with Rww' , $\mathfrak{M}, w' \Vdash \varphi$. Again by definition of $\mathfrak{M}, w \Vdash$, that is the case iff $\mathfrak{M}, w \Vdash \Diamond\varphi$. □

Problem 1.2. Complete the proof of [Proposition 1.7](#).

Problem 1.3. Let $\mathfrak{M} = \langle W, R, V \rangle$ be a model, and suppose $w_1, w_2 \in W$ are such that:

1. $w_1 \in V(p)$ if and only if $w_2 \in V(p)$; and
2. for all $w \in W$: Rw_1w if and only if Rw_2w .

Using induction on [formulas](#), show that for all [formulas](#) φ : $\mathfrak{M}, w_1 \Vdash \varphi$ if and only if $\mathfrak{M}, w_2 \Vdash \varphi$.

Problem 1.4. Let $\mathfrak{M} = \langle M, R, V \rangle$. Show that $\mathfrak{M}, w \Vdash \neg\Diamond\varphi$ if and only if $\mathfrak{M}, w \Vdash \Box\neg\varphi$.

1.6 Truth in a Model

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sec

Sometimes we are interested which [formulas](#) are true at every world in a given model. Let's introduce a notation for this.

Definition 1.8. A [formula](#) φ is *true in a model* $M = \langle W, R, V \rangle$, written $\mathfrak{M} \Vdash \varphi$, if and only if $\mathfrak{M}, w \Vdash \varphi$ for every $w \in W$.

mod:syn:tru:
prop:truthfacts

Proposition 1.9.

1. If $\mathfrak{M} \Vdash \varphi$ then $\mathfrak{M} \not\Vdash \neg\varphi$, but not vice-versa.
2. If $\mathfrak{M} \Vdash \varphi \rightarrow \psi$ then $\mathfrak{M} \Vdash \varphi$ only if $\mathfrak{M} \Vdash \psi$, but not vice-versa.

Proof. 1. If $\mathfrak{M} \Vdash \varphi$ then φ is true at all worlds in W , and since $W \neq \emptyset$, it can't be that $\mathfrak{M} \Vdash \neg\varphi$, or else φ would have to be both true and false at some world.

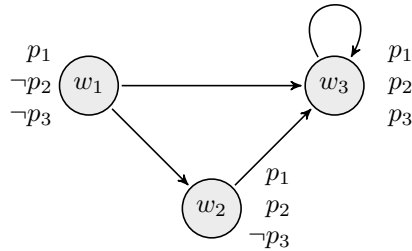
On the other hand, if $\mathfrak{M} \not\Vdash \neg\varphi$ then φ is true at some world $w \in W$. It does not follow that $\mathfrak{M}, w \Vdash \varphi$ for every $w \in W$. For instance, in the model of [Figure 1.1](#), $\mathfrak{M} \not\Vdash \neg p$, and also $\mathfrak{M} \not\Vdash p$.

- Assume $\mathfrak{M} \Vdash \varphi \rightarrow \psi$ and $\mathfrak{M} \Vdash \varphi$; to show $\mathfrak{M} \Vdash \psi$ let $w \in W$ be an arbitrary world. Then $\mathfrak{M}, w \Vdash \varphi \rightarrow \psi$ and $\mathfrak{M}, w \Vdash \varphi$, so $\mathfrak{M}, w \Vdash \psi$, and since w was arbitrary, $\mathfrak{M} \Vdash \psi$.

To show that the converse fails, we need to find a model \mathfrak{M} such that $\mathfrak{M} \Vdash \varphi$ only if $\mathfrak{M} \Vdash \psi$, but $\mathfrak{M} \not\Vdash \varphi \rightarrow \psi$. Consider again the model of [Figure 1.1](#): $\mathfrak{M} \not\Vdash p$ and hence (vacuously) $\mathfrak{M} \Vdash p$ only if $\mathfrak{M} \Vdash q$. However, $\mathfrak{M} \not\Vdash p \rightarrow q$, as p is true but q false at w_1 .

□

Problem 1.5. Consider the following model \mathfrak{M} for the language comprising p_1, p_2, p_3 as the only propositional variables:



Are the following **formulas** and schemas true in the model \mathfrak{M} , i.e., true at every world in \mathfrak{M} ? Explain.

- $p \rightarrow \Diamond p$ (for p atomic);
- $\varphi \rightarrow \Diamond \varphi$ (for φ arbitrary);
- $\Box p \rightarrow p$ (for p atomic);
- $\neg p \rightarrow \Diamond \Box p$ (for p atomic);
- $\Diamond \Box \varphi$ (for φ arbitrary);
- $\Box \Diamond p$ (for p atomic).

1.7 Validity

explanation

Formulas that are true in all models, i.e., true at every world in every model, are particularly interesting. They represent those modal propositions which are true regardless of how \Box and \Diamond are interpreted, as long as the interpretation is “normal” in the sense that it is generated by some accessibility relation on possible worlds. We call such **formulas** *valid*. For instance, $\Box(p \wedge q) \rightarrow \Box p$ is valid. Some **formulas** one might expect to be valid on the basis of the alethic interpretation of \Box , such as $\Box p \rightarrow p$, are not valid, however. Part of the interest of relational models is that different interpretations of \Box and \Diamond can be captured by different kinds of accessibility relations. This suggests that we should define validity not just relative to *all* models, but relative to all models *of a certain*

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sec

kind. It will turn out, e.g., that $\Box p \rightarrow p$ is true in all models where every world is accessible from itself, i.e., R is reflexive. Defining validity relative to classes of models enables us to formulate this succinctly: $\Box p \rightarrow p$ is valid in the class of reflexive models.

Definition 1.10. A formula φ is *valid* in a class \mathcal{C} of models if it is true in every model in \mathcal{C} (i.e., true at every world in every model in \mathcal{C}). If φ is valid in \mathcal{C} , we write $\mathcal{C} \models \varphi$, and we write $\models \varphi$ if φ is valid in the class of *all* models.

Proposition 1.11. *If φ is valid in \mathcal{C} it is also valid in each class $\mathcal{C}' \subseteq \mathcal{C}$.*

*mod:syn:val:
prop:Nec-rule*

Proposition 1.12. *If φ is valid, then so is $\Box\varphi$.*

Proof. Assume $\models \varphi$. To show $\models \Box\varphi$ let $\mathfrak{M} = \langle W, R, V \rangle$ be a model and $w \in W$. If Rww' then $\mathfrak{M}, w' \Vdash \varphi$, since φ is valid, and so also $\mathfrak{M}, w \Vdash \Box\varphi$. Since \mathfrak{M} and w were arbitrary, $\models \Box\varphi$. \square

Problem 1.6. Show that the following are valid:

1. $\models \Box p \rightarrow \Box(q \rightarrow p)$;
2. $\models \Box \neg \perp$;
3. $\models \Box p \rightarrow (\Box q \rightarrow \Box p)$.

Problem 1.7. Show that $\varphi \rightarrow \Box\varphi$ is valid in the class \mathcal{C} of models $\mathfrak{M} = \langle W, R, V \rangle$ where $W = \{w\}$. Similarly, show that $\psi \rightarrow \Box\psi$ and $\Diamond\varphi \rightarrow \psi$ are valid in the class of models $\mathfrak{M} = \langle W, R, V \rangle$ where $R = \emptyset$.

1.8 Tautological Instances

*mod:syn:tau:
sec*

A modal-free formula is a tautology if it is true under every truth-value assignment. Clearly, every tautology is true at every world in every model. But for formulas involving \Box and \Diamond , the notion of tautology is not defined. Is it the case, e.g., that $\Box p \vee \neg\Box p$ —an instance of the principle of excluded middle—is valid? The notion of a *tautological instance* helps: a formula that is a substitution instance of a (non-modal) tautology. It is not surprising, but still requires proof, that every tautological instance is valid.

explanation

Definition 1.13. A modal formula ψ is a *tautological instance* if and only if there is a modal-free tautology φ with propositional variables p_1, \dots, p_n and formulas $\theta_1, \dots, \theta_n$ such that $\psi \equiv \varphi[\theta_1/p_1, \dots, \theta_n/p_n]$.

*mod:syn:tau:
lem:valid-taut*

Lemma 1.14. *Suppose φ is a modal-free formula whose propositional variables are p_1, \dots, p_n , and let $\theta_1, \dots, \theta_n$ be modal formulas. Then for any assignment \mathbf{v} , any model $\mathfrak{M} = \langle W, R, V \rangle$, and any $w \in W$ such that $\mathbf{v}(p_i) = \mathbb{T}$ if and only if $\mathfrak{M}, w \Vdash \theta_i$ we have that $\mathbf{v} \models \varphi$ if and only if $\mathfrak{M}, w \Vdash \varphi[\theta_1/p_1, \dots, \theta_n/p_n]$.*

Proof. By induction on φ .

1. $\varphi \equiv \perp$: Both $\mathfrak{v} \not\models \perp$ and $\mathfrak{M}, w \not\vdash \perp$.
2. $\varphi \equiv \top$: Both $\mathfrak{v} \models \top$ and $\mathfrak{M}, w \vdash \top$.
3. $\varphi \equiv p_i$:

$\mathfrak{v} \models p_i \Leftrightarrow \mathfrak{v}(p_i) = \mathbb{T}$	by definition of $\mathfrak{v} \models p_i$;
$\Leftrightarrow \mathfrak{M}, w \Vdash \theta_i$	by assumption
$\Leftrightarrow \mathfrak{M}, w \Vdash p_i[\theta_1/p_1, \dots, \theta_n/p_n]$	since $p_i[\theta_1/p_1, \dots, \theta_n/p_n] \equiv \theta_i$.
4. $\varphi \equiv \neg\psi$:

$\mathfrak{v} \models \neg\psi \Leftrightarrow \mathfrak{v} \not\models \psi$	by definition of $\mathfrak{v} \models$;
$\Leftrightarrow \mathfrak{M}, w \not\vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$	by induction hypothesis;
$\Leftrightarrow \mathfrak{M}, w \Vdash \neg\psi[\theta_1/p_1, \dots, \theta_n/p_n]$	by definition of $\mathfrak{v} \models$.
5. $\varphi \equiv (\psi \wedge \chi)$:

$\mathfrak{v} \models \psi \wedge \chi \Leftrightarrow \mathfrak{v} \models \psi$ and $\mathfrak{v} \models \chi$	by definition of $\mathfrak{v} \models$;
$\Leftrightarrow \mathfrak{M}, w \Vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$ and	
$\mathfrak{M}, w \Vdash \chi[\theta_1/p_1, \dots, \theta_n/p_n]$,	by induction hypothesis;
$\Leftrightarrow \mathfrak{M}, w \Vdash (\psi \wedge \chi)[\theta_1/p_1, \dots, \theta_n/p_n]$	by definition of $\mathfrak{M}, w \Vdash$.
6. $\varphi \equiv (\psi \vee \chi)$:

$\mathfrak{v} \models \psi \vee \chi \Leftrightarrow \mathfrak{v} \models \psi$ or $\mathfrak{v} \models \chi$	by definition of $\mathfrak{v} \models$;
$\Leftrightarrow \mathfrak{M}, w \Vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$ or	
$\mathfrak{M}, w \Vdash \chi[\theta_1/p_1, \dots, \theta_n/p_n]$,	by induction hypothesis;
$\Leftrightarrow \mathfrak{M}, w \Vdash (\psi \vee \chi)[\theta_1/p_1, \dots, \theta_n/p_n]$	by definition of $\mathfrak{M}, w \Vdash$.
7. $\varphi \equiv (\psi \rightarrow \chi)$:

$\mathfrak{v} \models \psi \rightarrow \chi \Leftrightarrow \mathfrak{v} \not\models \psi$ or $\mathfrak{v} \models \chi$	by definition of $\mathfrak{v} \models$;
$\Leftrightarrow \mathfrak{M}, w \not\vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$ or	
$\mathfrak{M}, w \Vdash \chi[\theta_1/p_1, \dots, \theta_n/p_n]$,	by induction hypothesis;
$\Leftrightarrow \mathfrak{M}, w \Vdash (\psi \rightarrow \chi)[\theta_1/p_1, \dots, \theta_n/p_n]$	by definition of $\mathfrak{M}, w \Vdash$.
8. $\varphi \equiv (\psi \leftrightarrow \chi)$:

$\mathfrak{v} \models \psi \leftrightarrow \chi \Leftrightarrow$ either $\mathfrak{v} \models \psi$ and $\mathfrak{v} \models \chi$	
or $\mathfrak{v} \not\models \psi$ and $\mathfrak{v} \not\models \chi$	by definition of $\mathfrak{v} \models$;
\Leftrightarrow either $\mathfrak{M}, w \Vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$ and	
$\mathfrak{M}, w \Vdash \chi[\theta_1/p_1, \dots, \theta_n/p_n]$	
or $\mathfrak{M}, w \not\vdash \psi[\theta_1/p_1, \dots, \theta_n/p_n]$ and	
$\mathfrak{M}, w \not\vdash \chi[\theta_1/p_1, \dots, \theta_n/p_n]$,	by induction hypothesis;
$\Leftrightarrow \mathfrak{M}, w \Vdash (\psi \leftrightarrow \chi)[\theta_1/p_1, \dots, \theta_n/p_n]$	by definition of $\mathfrak{M}, w \Vdash$.

□

mod:syn:tau:
prop:valid-taut

Proposition 1.15. *All tautological instances are valid.*

Proof. Contrapositively, suppose φ is such that $\mathfrak{M}, w \not\models \varphi[\theta_1/p_1, \dots, \theta_n/p_n]$, for some model \mathfrak{M} and world w . Define an assignment v such that $v(p_i) = \mathbb{T}$ if and only if $\mathfrak{M}, w \models \theta_i$ (and v assigns arbitrary values to $q \notin \{p_1, \dots, p_n\}$). Then by Lemma 1.14, $\mathfrak{v} \not\models \varphi$, so φ is not a tautology. □

1.9 Schemas and Validity

mod:syn:sch:
sec

Definition 1.16. A *schema* is a set of formulas comprising all and only the substitution instances of some modal formula χ , i.e.,

$$\{\psi : \exists\theta_1, \dots, \exists\theta_n (\psi = \chi[\theta_1/p_1, \dots, \theta_n/p_n])\}.$$

The formula χ is called the *characteristic formula* of the schema, and it is unique up to a renaming of the propositional variables. A formula φ is an *instance* of a schema if it is a member of the set.

It is convenient to denote a schema by the meta-linguistic expression obtained by substituting ‘ φ ’, ‘ ψ ’, \dots , for the atomic components of χ . So, for instance, the following denote schemas: ‘ φ ’, ‘ $\varphi \rightarrow \Box\varphi$ ’, ‘ $\varphi \rightarrow (\psi \rightarrow \varphi)$ ’. They correspond to the characteristic formulas p , $p \rightarrow \Box p$, $p \rightarrow (q \rightarrow p)$. The schema ‘ φ ’ denotes the set of all formulas.

Definition 1.17. A schema is *true* in a model if and only if all of its instances are; and a schema is *valid* if and only if it is true in every model.

mod:syn:sch:
prop:Kvalid

Proposition 1.18. *The following schema K is valid*

$$\Box(\varphi \rightarrow \psi) \rightarrow (\Box\varphi \rightarrow \Box\psi). \quad (\text{K})$$

Proof. We need to show that all instances of the schema are true at every world in every model. So let $\mathfrak{M} = \langle W, R, V \rangle$ and $w \in W$ be arbitrary. To show that a conditional is true at a world we assume the antecedent is true to show that consequent is true as well. In this case, let $\mathfrak{M}, w \models \Box(\varphi \rightarrow \psi)$ and $\mathfrak{M}, w \models \Box\varphi$. We need to show $\mathfrak{M} \models \Box\psi$. So let w' be arbitrary such that Rww' . Then by the first assumption $\mathfrak{M}, w' \models \varphi \rightarrow \psi$ and by the second assumption $\mathfrak{M}, w' \models \varphi$. It follows that $\mathfrak{M}, w' \models \psi$. Since w' was arbitrary, $\mathfrak{M}, w \models \Box\psi$. □

mod:syn:sch:
prop:Dual-valid

Proposition 1.19. *The following schema DUAL is valid*

$$\Diamond\varphi \leftrightarrow \neg\Box\neg\varphi. \quad (\text{DUAL})$$

Proof. Exercise. □

Problem 1.8. Prove Proposition 1.19.

Valid Schemas	Invalid Schemas
$\Box(\varphi \rightarrow \psi) \rightarrow (\Diamond\varphi \rightarrow \Diamond\psi)$	$\Box(\varphi \vee \psi) \rightarrow (\Box\varphi \vee \Box\psi)$
$\Diamond(\varphi \rightarrow \psi) \rightarrow (\Box\varphi \rightarrow \Diamond\psi)$	$(\Diamond\varphi \wedge \Diamond\psi) \rightarrow \Diamond(\varphi \wedge \psi)$
$\Box(\varphi \wedge \psi) \leftrightarrow (\Box\varphi \wedge \Box\psi)$	$\varphi \rightarrow \Box\varphi$
$\Box\varphi \rightarrow \Box(\psi \rightarrow \varphi)$	$\Box\Diamond\varphi \rightarrow \psi$
$\neg\Diamond\varphi \rightarrow \Box(\varphi \rightarrow \psi)$	$\Box\Box\varphi \rightarrow \Box\varphi$
$\Diamond(\varphi \vee \psi) \leftrightarrow (\Diamond\varphi \vee \Diamond\psi)$	$\Box\Diamond\varphi \rightarrow \Diamond\Box\varphi$

Table 1.1: Valid and (or?) invalid schemas.

Proposition 1.20. *If φ and $\varphi \rightarrow \psi$ are true at a world in a model then so is ψ . Hence, the valid formulas are closed under modus ponens.*

mod:syn:sch:
tab:valid-invalidSchemas
mod:syn:sch:
prop:soundMP

Proposition 1.21. *A formula φ is valid iff all its substitution instances are. In other words, a schema is valid iff its characteristic formula is.*

mod:syn:sch:
prop:valid-instances

Proof. The “if” direction is obvious, since φ is a substitution instance of itself.

To prove the “only if” direction, we show the following: Suppose $\mathfrak{M} = \langle W, R, V \rangle$ is a modal model, and $\psi \equiv \varphi[D_1/p_1, \dots, D_n/p_n]$ is a substitution instance of φ . Define $\mathfrak{M}' = \langle W, R, V' \rangle$ by $V(p_i) = \{w : \mathfrak{M} \Vdash \theta_i w\}$. Then $\mathfrak{M} \Vdash \psi w$ iff $\mathfrak{M}' \Vdash \varphi w$, for any $w \in W$. (We leave the proof as an exercise.) Now suppose that φ was valid, but some substitution instance ψ of φ was not valid. Then for some $\mathfrak{M} = \langle W, R, V \rangle$ and some $w \in W$, $\mathfrak{M} \not\Vdash \psi w$. But then $\mathfrak{M}' \not\Vdash \varphi w$ by the claim, and φ is not valid, a contradiction. \square

Problem 1.9. Prove the claim in the “only if” part of the proof of Proposition 1.21. (Hint: use induction on φ .)

Note, however, that it is not true that a schema is true in a model iff its characteristic formula is. Of course, the “only if” direction holds: if every instance of φ is true in \mathfrak{M} , φ itself is true in \mathfrak{M} . But it may happen that φ is true in \mathfrak{M} but some instance of φ is false at some world in \mathfrak{M} . For a very simple counterexample consider p in a model with only one world w and $V(p) = \{w\}$, so that p is true at w . But \perp is an instance of p , and not true at w .

Problem 1.10. Show that none of the following formulas are valid:

- D: $\Box p \rightarrow \Diamond p$;
- T: $\Box p \rightarrow p$;
- B: $p \rightarrow \Box\Diamond p$;
- 4: $\Box p \rightarrow \Box\Box p$;
- 5: $\Diamond p \rightarrow \Box\Diamond p$.

Problem 1.11. Prove that the schemas in the first column of table 1.1 are valid and those in the second column are not valid.

Problem 1.12. Decide whether the following schemas are valid or invalid:

1. $(\Diamond\varphi \rightarrow \Box\psi) \rightarrow (\Box\varphi \rightarrow \Box\psi)$;
2. $\Diamond(\varphi \rightarrow \psi) \vee \Box(\psi \rightarrow \varphi)$.

Problem 1.13. For each of the following schemas find a model \mathfrak{M} such that every instance of the formula is true in \mathfrak{M} :

1. $p \rightarrow \Diamond\Diamond p$;
2. $\Diamond p \rightarrow \Box p$.

1.10 Entailment

mod:syn:ent:
sec

With the definition of truth at a world, we can define an entailment relation between formulas. A formula ψ entails φ iff, whenever ψ is true, φ is true as well. Here, “whenever” means both “whichever model we consider” as well as “whichever world in that model we consider.”

explanation

Definition 1.22. If Γ is a set of formulas and φ a formula, then Γ entails φ , in symbols: $\Gamma \vDash \varphi$, if and only if for every model $\mathfrak{M} = \langle W, R, V \rangle$ and world $w \in W$, if $\mathfrak{M}, w \Vdash \psi$ for every $\psi \in \Gamma$, then $\mathfrak{M}, w \Vdash \varphi$. If Γ contains a single formula ψ , then we write $\psi \vDash \varphi$.

Example 1.23. To show that a formula entails another, we have to reason about all models, using the definition of $\mathfrak{M}, w \Vdash$. For instance, to show $p \rightarrow \Diamond p \vDash \Box \neg p \rightarrow \neg p$, we might argue as follows: Consider a model $\mathfrak{M} = \langle W, R, V \rangle$ and $w \in W$, and suppose $\mathfrak{M}, w \Vdash p \rightarrow \Diamond p$. We have to show that $\mathfrak{M}, w \Vdash \Box \neg p \rightarrow \neg p$. Suppose not. Then $\mathfrak{M}, w \Vdash \Box \neg p$ and $\mathfrak{M}, w \not\Vdash \neg p$. Since $\mathfrak{M}, w \not\Vdash \neg p$, $\mathfrak{M}, w \Vdash p$. By assumption, $\mathfrak{M}, w \Vdash p \rightarrow \Diamond p$, hence $\mathfrak{M}, w \Vdash \Diamond p$. By definition of $\mathfrak{M}, w \Vdash \Diamond p$, there is some w' with Rww' such that $\mathfrak{M}, w' \Vdash p$. Since also $\mathfrak{M}, w \Vdash \Box \neg p$, $\mathfrak{M}, w' \Vdash \neg p$, a contradiction.

To show that a formula ψ does not entail another φ , we have to give a counterexample, i.e., a model $\mathfrak{M} = \langle W, R, V \rangle$ where we show that at some world $w \in W$, $\mathfrak{M}, w \Vdash \psi$ but $\mathfrak{M}, w \not\Vdash \varphi$. Let's show that $p \rightarrow \Diamond p \not\vDash \Box p \rightarrow p$. Consider the model in Figure 1.2. We have $\mathfrak{M}, w_1 \Vdash \Diamond p$ and hence $\mathfrak{M}, w_1 \Vdash p \rightarrow \Diamond p$. However, since $\mathfrak{M}, w_1 \Vdash \Box p$ but $\mathfrak{M}, w_1 \not\Vdash p$, we have $\mathfrak{M}, w_1 \not\vDash \Box p \rightarrow p$.

Often very simple counterexamples suffice. The model $\mathfrak{M}' = \{W', R', V'\}$ with $W' = \{w\}$, $R' = \emptyset$, and $V'(p) = \emptyset$ is also a counterexample: Since $\mathfrak{M}', w \not\vDash p$, $\mathfrak{M}', w \Vdash p \rightarrow \Diamond p$. As no worlds are accessible from w , we have $\mathfrak{M}', w \Vdash \Box p$, and so $\mathfrak{M}', w \not\vDash \Box p \rightarrow p$.

Problem 1.14. Show that $\Box(\varphi \wedge \psi) \vDash \Box\varphi$.

Problem 1.15. Show that $\Box(p \rightarrow q) \not\vDash p \rightarrow \Box q$ and $p \rightarrow \Box q \not\vDash \Box(p \rightarrow q)$.

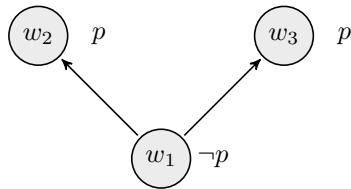


Figure 1.2: Counterexample to $p \rightarrow \Diamond p \models \Box p \rightarrow p$.

mod:syn:ent:
fig:counterex

Chapter 2

Frame Definability

2.1 Introduction

mod:fd:int:
sec

One question that interests modal logicians is the relationship between the accessibility relation and the truth of certain **formulas** in models with that accessibility relation. For instance, suppose the accessibility relation is reflexive, i.e., for every $w \in W$, Rww . In other words, every world is accessible from itself. That means that when $\Box\varphi$ is true at a world w , w itself is among the accessible worlds at which φ must therefore be true. So, if the accessibility relation R of \mathfrak{M} is reflexive, then whatever world w and formula φ we take, $\Box\varphi \rightarrow \varphi$ will be true there (in other words, the schema $\Box p \rightarrow p$ and all its substitution instances are true in \mathfrak{M}).

The converse, however, is false. It's not the case, e.g., that if $\Box p \rightarrow p$ is true in \mathfrak{M} , then R is reflexive. For we can easily find a non-reflexive model \mathfrak{M} where $\Box p \rightarrow p$ is true at all worlds: take the model with a single world w , not accessible from itself, but with $w \in V(p)$. By picking the truth value of p suitably, we can make $\Box\varphi \rightarrow \varphi$ true in a model that is not reflexive.

The solution is to remove the variable assignment V from the equation. If we require that $\Box p \rightarrow p$ is true at all worlds in \mathfrak{M} , regardless of which worlds are in $V(p)$, then it is necessary that R is reflexive. For in any non-reflexive model, there will be at least one world w such that not Rww . If we set $V(p) = W \setminus \{w\}$, then p will be true at all worlds other than w , and so at all worlds accessible from w (since w is guaranteed not to be accessible from w , and w is the only world where p is false). On the other hand, p is false at w , so $\Box p \rightarrow p$ is false at w .

This suggests that we should introduce a notation for model structures without a valuation: we call these *frames*. A frame \mathfrak{F} is simply a pair $\langle W, R \rangle$ consisting of a set of worlds with an accessibility relation. Every model $\langle W, R, V \rangle$ is then, as we say, *based on* the frame $\langle W, R \rangle$. Conversely, a frame determines the class of models based on it; and a class of frames determines the class of models which are based on any frame in the class. And we can define $\mathfrak{F} \models \varphi$, the notion of a **formula** being *valid* in a frame as: $\mathfrak{M} \Vdash \varphi$ for all \mathfrak{M} based on \mathfrak{F} .

If R is ...	then ... is true in \mathfrak{M} :
<i>serial</i> : $\forall u \exists v Ruv$	$\Box p \rightarrow \Diamond p$ (D)
<i>reflexive</i> : $\forall w Rww$	$\Box p \rightarrow p$ (T)
<i>symmetric</i> : $\forall u \forall v (Ruv \rightarrow Rvu)$	$p \rightarrow \Box \Diamond p$ (B)
<i>transitive</i> : $\forall u \forall v \forall w ((Ruv \wedge Rvw) \rightarrow Ruw)$	$\Box p \rightarrow \Box \Box p$ (4)
<i>euclydean</i> : $\forall w \forall u \forall v ((Rwu \wedge Rvw) \rightarrow Ruw)$	$\Diamond p \rightarrow \Box \Diamond p$ (5)

Table 2.1: Five correspondence facts.

With this notation, we can establish correspondence relations between **formulas** and classes of frames: e.g., $\mathfrak{F} \models \Box p \rightarrow p$ if, and only if, \mathfrak{F} is reflexive.

mod:frd:acc:
tab:five

2.2 Properties of Accessibility Relations

Many modal **formulas** turn out to be characteristic of simple, and even familiar, properties of the accessibility relation. In one direction, that means that any model that has a given property makes a corresponding **formula** (and all its substitution instances) true. We begin with five classical examples of kinds of accessibility relations and the formulas the truth of which they guarantee.

mod:frd:acc:
sec

Theorem 2.1. *Let $\mathfrak{M} = \langle W, R, V \rangle$ be a model. If R has the property on the left side of table 2.1, every instance of the **formula** on the right side is true in \mathfrak{M} .*

mod:frd:acc:
thm:soundschemas

Proof. Here is the case for B: to show that the schema is true in a model we need to show that all of its instances are true all worlds in the model. So let $\varphi \rightarrow \Box \Diamond \varphi$ be a given instance of B, and let $w \in W$ be an arbitrary world. Suppose the antecedent φ is true at w , in order to show that $\Box \Diamond \varphi$ is true at w . So we need to show that $\Diamond \varphi$ is true at all w' accessible from w . Now, for any w' such that Rww' we have, using the hypothesis of symmetry, that also $Rw'w$ (see Figure 2.1). Since $\mathfrak{M}, w \models \varphi$, we have $\mathfrak{M}, w' \models \Diamond \varphi$. Since w' was an arbitrary world such that Rww' , we have $\mathfrak{M}, w \models \Box \Diamond \varphi$.

We leave the other cases as exercises. □

Problem 2.1. Complete the proof of Theorem 2.1

Notice that the converse implications of Theorem 2.1 do not hold: it's not true that if a model verifies a schema, then the accessibility relation of that model has the corresponding property. In the case of T and reflexive models, it is easy to give an example of a model in which T itself fails: let $W = \{w\}$ and $V(p) = \emptyset$. Then R is not reflexive, but $\mathfrak{M}, w \models \Box p$ and $\mathfrak{M}, w \not\models p$. But here we have just a single instance of T that fails in \mathfrak{M} , other instances, e.g., $\Box \neg p \rightarrow \neg p$

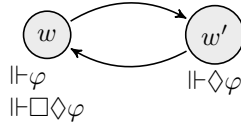


Figure 2.1: The argument from symmetry.

mod:frd:acc:
fig:Bsymm

are true. It is harder to give examples where *every substitution instance* of T is true in \mathfrak{M} and \mathfrak{M} is not reflexive. But there are such models, too:

mod:frd:acc:
prop:reflexive

Proposition 2.2. *Let $\mathfrak{M} = \langle W, R, V \rangle$ be a model such that $W = \{u, v\}$, where worlds u and v are related by R : i.e., both Ruv and Rvu . Suppose that for all p : $u \in V(p) \Leftrightarrow v \in V(p)$. Then:*

1. *For all φ : $\mathfrak{M}, u \Vdash \varphi$ if and only if $\mathfrak{M}, v \Vdash \varphi$ (use induction on φ).*
2. *Every instance of T is true in \mathfrak{M} .*

Since \mathfrak{M} is not reflexive (it is, in fact, irreflexive), the converse of [Theorem 2.1](#) fails in the case of T (similar arguments can be given for some—though not all—the other schemas mentioned in [Theorem 2.1](#)).

Problem 2.2. Prove the claims in [Proposition 2.2](#).

Although we will focus on the five classical [formulas](#) D, T, B, 4, and 5, we record in [table 2.2](#) a few more properties of accessibility relations. The accessibility relation R is partially functional, if from every world at most one world is accessible. If it is the case that from every world exactly one world is accessible, we call it functional. (Thus the functional relations are precisely those that are both serial and partially functional). They are called “functional” because the accessibility relation operates like a (partial) function. A relation is weakly dense if whenever Ruv , there is a w “between” u and v . So weakly dense relations are in a sense the opposite of transitive relations: in a transitive relation, whenever you can reach v from u by a detour via w , you can reach v from u directly; in a weakly dense relation, whenever you can reach v from u directly, you can also reach it by a detour via some w . A relation is weakly directed if whenever you can reach worlds u and v from some world w , you can reach a single world t from both u and v —this is sometimes called the “diamond property” or “confluence.”

Problem 2.3. Let $\mathfrak{M} = \langle W, R, V \rangle$ be a model. Show that if R satisfies the left-hand properties of [table 2.2](#), every instance of the corresponding right-hand formula is true in \mathfrak{M} .

2.3 Frames

mod:frd:fra:
sec

If R is ...	then ... is true in \mathfrak{M} :
<i>partially functional</i> : $\forall w \forall u \forall v ((Rwu \wedge Rvw) \rightarrow u = v)$	$\Diamond p \rightarrow \Box p$
<i>functional</i> : $\forall w \exists v \forall u (Rwu \leftrightarrow u = v)$	$\Diamond p \leftrightarrow \Box p$
<i>weakly dense</i> : $\forall u \forall v (Ruv \rightarrow \exists w (Ruw \wedge Rvw))$	$\Box \Box p \rightarrow \Box p$
<i>weakly connected</i> : $\forall w \forall u \forall v ((Rwu \wedge Rvw) \rightarrow (Ruv \vee u = v \vee Rvu))$	$\Box((p \wedge \Box p) \rightarrow q) \vee \Box((q \wedge \Box q) \rightarrow p)$ (L)
<i>weakly directed</i> : $\forall w \forall u \forall v ((Rwu \wedge Rvw) \rightarrow \exists t (Rut \wedge Rvt))$	$\Diamond \Box p \rightarrow \Box \Diamond p$ (G)

Table 2.2: Five more correspondence facts.

mod:frd:acc:
tab:anotherfive

Definition 2.3. A *frame* is a pair $\mathfrak{F} = \langle W, R \rangle$ where W is a non-empty set of worlds and R a binary relation on W . A model \mathfrak{M} is *based on* a frame $\mathfrak{F} = \langle W, R \rangle$ if and only if $\mathfrak{M} = \langle W, R, V \rangle$.

Definition 2.4. If \mathfrak{F} is a frame, we say that φ is *valid in* \mathfrak{F} , $\mathfrak{F} \vDash \varphi$, if $\mathfrak{M} \Vdash \varphi$ for every model \mathfrak{M} based on \mathfrak{F} .

If \mathcal{F} is a class of frames, we say φ is *valid in* \mathcal{F} , $\mathcal{F} \vDash \varphi$, iff $\mathfrak{F} \vDash \varphi$ for every frame $\mathfrak{F} \in \mathcal{F}$.

The reason frames are interesting is that correspondence between schemas and properties of the accessibility relation R is at the level of frames, *not of models*. For instance, although T is true in all reflexive models, not every model in which T is true is reflexive. However, it *is* true that not only is T *valid* on all reflexive *frames*, also every frame in which T is valid is reflexive.

Remark 1. Validity in a class of frames is a special case of the notion of validity in a class of models: $\mathcal{F} \vDash \varphi$ iff $\mathcal{C} \vDash \varphi$ where \mathcal{C} is the class of all models based on a frame in \mathcal{F} .

Obviously, if a *formula* or a schema is valid, i.e., valid with respect to the class of *all* models, it is also valid with respect to any class \mathcal{F} of frames.

2.4 Frame Definability

Even though the converse implications of [Theorem 2.1](#) fail, they hold if we replace “model” by “frame”: for the properties considered in [Theorem 2.1](#), it *is* true that if a *formula* is valid in a *frame* then the accessibility relation of that frame has the corresponding property. So, the *formulas* considered *define* the classes of frames that have the corresponding property.

mod:frd:def:
sec

Definition 2.5. If \mathcal{C} is a class of frames, we say φ *defines* \mathcal{C} iff $\mathfrak{F} \vDash \varphi$ for all and only frames $\mathfrak{F} \in \mathcal{C}$.

We now proceed to establish the full definability results for frames.

mod:frd:def: **Theorem 2.6.** *If the formula on the right side of table 2.1 is valid in a*
thm:fullCorrespondence *frame \mathfrak{F} , then \mathfrak{F} has the property on the left side.*

- Proof.* 1. Suppose D is valid in $\mathfrak{F} = \langle W, R \rangle$, i.e., $\mathfrak{F} \models \Box p \rightarrow \Diamond p$. Let $\mathfrak{M} = \langle W, R, V \rangle$ be a model based on \mathfrak{F} , and $w \in W$. We have to show that there is a v such that Rwv . Suppose not: then both $\mathfrak{M} \Vdash \Box \varphi$ and $\mathfrak{M}, w \not\Vdash \Diamond \varphi$ for any φ , including p . But then $\mathfrak{M}, w \not\Vdash \Box p \rightarrow \Diamond p$, contradicting the assumption that $\mathfrak{F} \models \Box p \rightarrow \Diamond p$.
2. Suppose T is valid in \mathfrak{F} , i.e., $\mathfrak{F} \models \Box p \rightarrow p$. Let $w \in W$ be an arbitrary world; we need to show Rww . Let $u \in V(p)$ if and only if Rwu (when q is other than p , $V(q)$ is arbitrary, say $V(q) = \emptyset$). Let $\mathfrak{M} = \langle W, R, V \rangle$. By construction, for all u such that Rwu : $\mathfrak{M}, u \Vdash p$, and hence $\mathfrak{M}, w \Vdash \Box p$. But by hypothesis $\Box p \rightarrow p$ is true at w , so that $\mathfrak{M}, w \Vdash p$, but by definition of V this is possible only if Rww .
3. We prove the contrapositive: Suppose \mathfrak{F} is not symmetric, we show that B, i.e., $p \rightarrow \Box \Diamond p$ is not valid in $\mathfrak{F} = \langle W, R \rangle$. If \mathfrak{F} is not symmetric, there are $u, v \in W$ such that Ruv but not Rvu . Define V such that $w \in V(p)$ if and only if not Rvw (and V is arbitrary otherwise). Let $\mathfrak{M} = \langle W, R, V \rangle$. Now, by definition of V , $\mathfrak{M}, w \Vdash p$ for all w such that not Rvw , in particular, $\mathfrak{M}, u \Vdash p$ since not Rvu . Also, since Rvw iff $p \notin V(w)$, there is no w such that Rvw and $\mathfrak{M}, w \Vdash p$, and hence $\mathfrak{M}, v \not\Vdash \Diamond p$. Since Ruv , also $\mathfrak{M}, u \not\Vdash \Box \Diamond p$. It follows that $\mathfrak{M}, u \not\Vdash p \rightarrow \Box \Diamond p$, and so B is not valid in \mathfrak{F} .
4. Suppose 4 is valid in $\mathfrak{F} = \langle W, R \rangle$, i.e., $\mathfrak{F} \models \Box p \rightarrow \Box \Box p$, and let $u, v, w \in W$ be arbitrary worlds such that Ruw and Rvw ; we need to show that Ruv . Define V such that $z \in V(p)$ if and only if Ruz (and V is arbitrary otherwise). Let $\mathfrak{M} = \langle W, R, V \rangle$. By definition of V , $\mathfrak{M}, z \Vdash p$ for all z such that Ruz , and hence $\mathfrak{M}, u \Vdash \Box p$. But by hypothesis 4, $\Box p \rightarrow \Box \Box p$, is true at u , so that $\mathfrak{M}, u \Vdash \Box \Box p$. Since Ruw and Rvw , we have $\mathfrak{M}, w \Vdash p$, but by definition of V this is possible only if Ruv , as desired.
5. We proceed contrapositively, assuming that the frame $\mathfrak{F} = \langle W, R \rangle$ is not euclidean, and show that it falsifies 5, i.e., i.e., $\mathfrak{F} \not\models \Diamond p \rightarrow \Box \Diamond p$. Suppose there are worlds $u, v, w \in W$ such that Rwu and Rvw but not Ruv . Define V such that for all worlds z , $z \in V(p)$ if and only if it is *not* the case that Ruz . Let $\mathfrak{M} = \langle W, R, V \rangle$. Then by hypothesis $\mathfrak{M}, v \Vdash p$ and since Rvw also $\mathfrak{M}, w \Vdash \Diamond p$. However, there is no world y such that Ruy and $\mathfrak{M}, y \Vdash p$ so $\mathfrak{M}, u \not\Vdash \Diamond p$. Since Rwu , it follows that $\mathfrak{M}, w \not\Vdash \Box \Diamond p$, so that 5, $\Diamond p \rightarrow \Box \Diamond p$, fails at w .

□

You'll notice a difference between the proof for D and the other cases: no mention was made of the valuation V . In effect, we proved that if $\mathfrak{M} \Vdash D$ then \mathfrak{M} is serial. So D defines the class of serial *models*, not just frames.

Corollary 2.7. *Any model where D is true is serial.*

*mod:frd:def:
prop:D-serial*

Corollary 2.8. *Each formula on the right side of table 2.1 defines the class of frames which have the property on the left side.*

Proof. In Theorem 2.1, we proved that if a model has the property on the left, the formula on the right is true in it. Thus, if a frame \mathfrak{F} has the property on the left, the formula on the left is valid in \mathfrak{F} . In Theorem 2.6, we proved the converse implications: if a formula on the right is valid in \mathfrak{F} , \mathfrak{F} has the property on the left. \square

Problem 2.4. Show that if the formula on the right side of table 2.2 is valid in a frame \mathfrak{F} , then \mathfrak{F} has the property on the left side. To do this, consider a frame that does *not* satisfy the property on the left, and define a suitable V such that the formula on the right is false at some world.

Theorem 2.6 also shows that the properties can be combined: for instance if both B and 4 are valid in \mathfrak{F} then the frame is both symmetric and transitive, etc. Many important modal logics are characterized as the set of formulas valid in all frames that combine some frame properties, and so we can characterize them as the set of formulas valid in all frames in which the corresponding defining formulas are valid. For instance, the classical system **S4** is the set of all formulas valid in all reflexive and transitive frames, i.e., in all those where both T and 4 are valid. **S5** is the set of all formulas valid in all reflexive, symmetric, and euclidean frames, i.e., all those where all of T, B, and 5 are valid.

Logical relationships between properties of R in general correspond to relationships between the corresponding defining formulas. For instance, every reflexive relation is serial; hence, whenever T is valid in a frame, so is D. (Note that this relationship is *not* that of entailment. It is not the case that whenever $\mathfrak{M}, w \Vdash T$ then $\mathfrak{M}, w \Vdash D$.) We record some such relationships.

Proposition 2.9. *Let R be a binary relation on a set W ; then:*

*mod:frd:def:
prop:relation-facts*

1. *If R is reflexive, then it is serial.*
2. *If R is symmetric, then it is transitive if and only if it is euclidean.*
3. *If R is symmetric or euclidean then it is weakly directed (it has the “diamond property”).*
4. *If R is euclidean then it is weakly connected.*
5. *If R is functional then it is serial.*

Problem 2.5. Prove Proposition 2.9.

2.5 First-order Definability

mod:frd:fol:
sec

We've seen that a number of properties of accessibility relations of frames can be defined by modal **formulas**. For instance, symmetry of frames can be defined by the **formula** $B, p \rightarrow \Box \Diamond p$. The conditions we've encountered so far can all be expressed by first-order **formulas** in a **language** involving a single two-place **predicate symbol**. For instance, symmetry is defined by $\forall x \forall y (Q(x, y) \rightarrow Q(y, x))$ in the sense that a first-order **structure** \mathfrak{M} with $|\mathfrak{M}| = W$ and $Q^{\mathfrak{M}} = R$ satisfies the preceding formula iff R is symmetric. This suggests the following definition:

Definition 2.10. A class \mathcal{C} of frames is *first-order definable* if there is a **sentence** φ in the first-order language with a single two-place **predicate symbol** P such that $\mathfrak{F} = \langle W, R \rangle \in \mathcal{C}$ iff $\mathfrak{M} \models \varphi$ in the first-order **structure** \mathfrak{M} with $|\mathfrak{M}| = W$ and $Q^{\mathfrak{M}} = R$.

It turns out that the properties and modal **formulas** that define them considered so far are exceptional. Not every **formula** defines a first-order definable class of frames, and not every first-order definable class of frames is definable by a modal formula.

A counterexample to the first is given by the Löb formula:

$$\Box(\Box p \rightarrow p) \rightarrow \Box p. \quad (\text{W})$$

W defines the class of transitive and converse well-founded frames. A relation is well-founded if there is no infinite sequence w_1, w_2, \dots such that Rw_2w_1, Rw_3w_2, \dots . For instance, the relation $<$ on \mathbb{N} is well-founded, whereas the relation $<$ on \mathbb{Z} is not. A relation is converse well-founded iff its converse is well-founded. So converse well-founded relations are those where there is no infinite sequence w_1, w_2, \dots such that Rw_1w_2, Rw_2w_3, \dots .

There is, however, no first-order **formula** defining transitive converse well-founded relations. For suppose $\mathfrak{M} \models \beta$ iff $R = Q^{\mathfrak{M}}$ is transitive converse well-founded. Let φ_n be the **formula**

$$(Q(a_1, a_2) \wedge \dots \wedge Q(a_{n-1}, a_n))$$

Now consider the set of **formulas**

$$\Gamma = \{\beta, \varphi_1, \varphi_2, \dots\}.$$

Every finite subset of Γ is satisfiable: Let k be largest such that φ_k is in the subset, $|\mathfrak{M}_k| = \{1, \dots, k\}$, $a_i^{\mathfrak{M}_k} = i$, and $P^{\mathfrak{M}_k} = <$. Since $<$ on $\{1, \dots, k\}$ is transitive and converse well-founded, $\mathfrak{M}_k \models \beta$. $\mathfrak{M}_k \models \varphi_i$ by construction, for all $i \leq k$. By the Compactness Theorem for first-order logic, Γ is satisfiable in some **structure** \mathfrak{M} . By hypothesis, since $\mathfrak{M} \models \beta$, the relation $Q^{\mathfrak{M}}$ is converse well-founded. But clearly, $a_1^{\mathfrak{M}}, a_2^{\mathfrak{M}}, \dots$ would form an infinite sequence of the kind ruled out by converse well-foundedness.

A counterexample to the second claim is given by the property of universality: for every u and v , Ruv . Universal frames are first-order definable by

the formula $\forall x \forall y Q(x, y)$. However, no modal formula is valid in all and only the universal frames. This is a consequence of a result that is independently interesting: the formulas valid in universal frames are exactly the same as those valid in reflexive, symmetric, and transitive frames. There are reflexive, symmetric, and transitive frames that are not universal, hence every formula valid in all universal frames is also valid in some non-universal frames.

2.6 Equivalence Relations and S5

The modal logic **S5** is characterized as the set of formulas valid on all universal frames, i.e., every world is accessible from every world, including itself. In such a scenario, \Box corresponds to necessity and \Diamond to possibility: $\Box\varphi$ is true if φ is true at every world, and $\Diamond\varphi$ is true if φ is true at some world. It turns out that **S5** can also be characterized as the formulas valid on all reflexive, symmetric, and transitive frames, i.e., on all *equivalence relations*.

mod:frd:es5:
sec

Definition 2.11. A binary relation R on W is an *equivalence relation* if and only if it is reflexive, symmetric and transitive. A relation R on W is *universal* if and only if Ruv for all $u, v \in W$.

Since T, B, and 4 characterize the reflexive, symmetric, and transitive frames, the frames where the accessibility relation is an equivalence relation are exactly those in which all four formulas are valid. It turns out that the equivalence relations can also be characterized by other combinations of formulas, since the conditions with which we've defined equivalence relations are equivalent to combinations of other familiar conditions on R .

Proposition 2.12. *The following are equivalent:*

mod:frd:es5:
prop:equivalences

1. R is an equivalence relation;
2. R is reflexive and euclidean;
3. R is serial, symmetric, and euclidean;
4. R is serial, symmetric, and transitive.

Proof. Exercise. □

Problem 2.6. Prove Proposition 2.12 by showing:

1. If R is symmetric and transitive, it is euclidean.
2. If R is reflexive, it is serial.
3. If R is reflexive and euclidean, it is symmetric.
4. If R is symmetric and euclidean, it is transitive.
5. If R is serial, symmetric, and transitive, it is reflexive.

Explain why this suffices for the proof that the conditions are equivalent.

Proposition 2.12 is the semantic counterpart to **Proposition 3.29**, in that it gives equivalent characterization of the modal logic of frames over which R is an equivalence (the logic traditionally referred to as **S5**).

What is the relationship between universal and equivalence relations? Although every universal relation is an equivalence relation, clearly not every equivalence relation is universal. However, the formulas valid on all universal relations are exactly the same as those valid on all equivalence relations.

Proposition 2.13. *Let R be an equivalence relation, and for each $w \in W$ define the equivalence class of w as the set $[w] = \{w' \in W : Rww'\}$. Then:*

1. $w \in [w]$;
2. R is universal on each equivalence class $[w]$;
3. The collection of equivalence classes partitions W into mutually exclusive and jointly exhaustive subsets.

*mod:frd:es5:
prop:S5=univ*

Proposition 2.14. *A formula φ is valid in all frames $\mathfrak{F} = \langle W, R \rangle$ where R is an equivalence relation, if and only if it is valid in all frames $\mathfrak{F} = \langle W, R \rangle$ where R is universal. Hence, the logic of universal frames is just **S5**.*

Proof. It's immediate to verify that a universal relation R on W is an equivalence. Hence, if φ is valid in all frames where R is an equivalence it is valid in all universal frames. For the other direction, we argue contrapositively: suppose ψ is a formula that fails at a world w in a model $\mathfrak{M} = \langle W, R, V \rangle$ based on a frame $\langle W, R \rangle$, where R is an equivalence on W . So $\mathfrak{M}, w \not\models \psi$. Define a model $\mathfrak{M}' = \langle W', R', V' \rangle$ as follows:

1. $W' = [w]$;
2. R' is universal on W' ;
3. $V'(p) = V(p) \cap W'$.

(So the set W' of worlds in \mathfrak{M}' is represented by the shaded area in **Figure 2.2**.) It is easy to see that R and R' agree on W' . Then one can show by induction on formulas that for all $w' \in W'$: $\mathfrak{M}', w' \models \varphi$ if and only if $\mathfrak{M}, w' \models \varphi$ for each φ (this makes sense since $W' \subseteq W$). In particular, $\mathfrak{M}', w \not\models \psi$, and ψ fails in a model based on a universal frame. \square

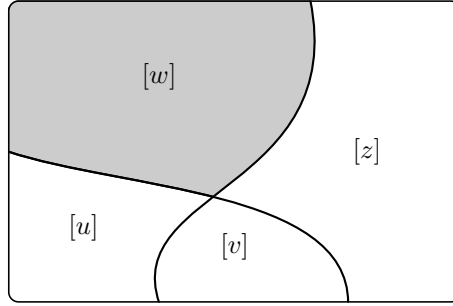


Figure 2.2: A partition of W in equivalence classes.

mod:frd:es5:
fig:partition

2.7 Second-order Definability

Not every frame property definable by modal formulas is first-order definable. However, if we allow quantification over one-place predicates (i.e., monadic second-order quantification), we define all modally definable frame properties. The trick is to exploit a systematic way in which the conditions under which a modal **formula** is true at a world are related to first-order **formulas**. This is the so-called standard translation of modal formulas into first-order formulas in a language containing not just a two-place **predicate symbol** Q for the accessibility relation, but also a one-place **predicate symbol** P_i for the **propositional variables** p_i occurring in φ .

mod:frd:st:
sec

Definition 2.15. The *standard translation* $ST_x(\varphi)$ is inductively defined as follows:

1. $\varphi \equiv \perp$: $ST_x(\varphi) = \perp$.
2. $\varphi \equiv \top$: $ST_x(\varphi) = \top$.
3. $\varphi \equiv p_i$: $ST_x(\varphi) = P_i(x)$.
4. $\varphi \equiv \neg\psi$: $ST_x(\varphi) = \neg ST_x(\psi)$.
5. $\varphi \equiv (\psi \wedge \chi)$: $ST_x(\varphi) = (ST_x(\psi) \wedge ST_x(\chi))$.
6. $\varphi \equiv (\psi \vee \chi)$: $ST_x(\varphi) = (ST_x(\psi) \vee ST_x(\chi))$.
7. $\varphi \equiv (\psi \rightarrow \chi)$: $ST_x(\varphi) = (ST_x(\psi) \rightarrow ST_x(\chi))$.
8. $\varphi \equiv (\psi \leftrightarrow \chi)$: $ST_x(\varphi) = (ST_x(\psi) \leftrightarrow ST_x(\chi))$.
9. $\varphi \equiv \Box\psi$: $ST_x(\varphi) = \forall y (Q(x, y) \rightarrow ST_y(\psi))$.
10. $\varphi \equiv \Diamond\psi$: $ST_x(\varphi) = \exists y (Q(x, y) \wedge ST_y(\psi))$.

For instance, $\text{ST}_x(\Box p \rightarrow p)$ is $\forall y (Q(x, y) \rightarrow P(y)) \rightarrow P(x)$. Any **structure** for the language of $\text{ST}_x(\varphi)$ requires a domain, a two-place relation assigned to Q , and subsets of the domain assigned to the one-place **predicate symbols** P_i . In other words, the components of such a **structure** are exactly those of a model for φ : the domain is the set of worlds, the two-place relation assigned to Q is the accessibility relation, and the subsets assigned to P_i are just the assignments $V(\rho_i)$. It won't surprise that satisfaction of φ in a modal model and of $\text{ST}_x(\varphi)$ in the corresponding **structure** agree:

mod:frd:st: prop:st **Proposition 2.16.** *Let $\mathfrak{M} = \langle W, R, V \rangle$, \mathfrak{M}' be the first-order **structure** with $|\mathfrak{M}'| = W$, $Q^{\mathfrak{M}'} = R$, and $P_i^{\mathfrak{M}'} = V(\rho_i)$, and $s(x) = w$. Then*

$$\mathfrak{M}, w \Vdash \varphi \text{ iff } \mathfrak{M}', s \models \text{ST}_x(\varphi)$$

Proof. By induction on φ . □

Proposition 2.17. *Suppose φ is a modal **formula** and $\mathfrak{F} = \langle W, R \rangle$ is a frame. Let \mathfrak{F}' be the first-order **structure** with $|\mathfrak{F}'| = W$ and $Q^{\mathfrak{F}'} = R$, and let φ' be the second-order **formula***

$$\forall X_1 \dots \forall X_n \forall x \text{ST}_x(\varphi)[X_1/P_1, \dots, X_n/P_n],$$

where P_1, \dots, P_n are all one-place **predicate symbols** in $\text{ST}_x(\varphi)$. Then

$$\mathfrak{F} \models \varphi \text{ iff } \mathfrak{F}' \models \varphi'$$

Proof. $\mathfrak{F}' \models \varphi'$ iff for every **structure** \mathfrak{M}' where $P_i^{\mathfrak{M}'} \subseteq W$ for $i = 1, \dots, n$, and for every s with $s(x) \in W$, $\mathfrak{M}', s \models \text{ST}_x(\varphi)$. By **Proposition 2.16**, that is the case iff for all models \mathfrak{M} based on \mathfrak{F} and every world $w \in W$, $\mathfrak{M}, w \Vdash \varphi$, i.e., $\mathfrak{F} \models \varphi$. □

Definition 2.18. A class \mathcal{C} of frames is *second-order definable* if there is a **sentence** φ in the second-order language with a single two-place **predicate symbol** P and quantifiers only over monadic set variables such that $\mathfrak{F} = \langle W, R \rangle \in \mathcal{C}$ iff $\mathfrak{M} \models \varphi$ in the **structure** \mathfrak{M} with $|\mathfrak{M}| = W$ and $P^{\mathfrak{M}} = R$.

Corollary 2.19. *If a class of frames is definable by a **formula** φ , the corresponding class of accessibility relations is definable by a monadic second-order **sentence**.*

Proof. The monadic second-order sentence φ' of the preceding proof has the required property. □

As an example, consider again the **formula** $\Box p \rightarrow p$. It defines reflexivity. Reflexivity is of course first-order definable by the **sentence** $Q(x, x)$. But it is also definable by the monadic second-order **sentence**

$$\forall X \forall x (\forall y (Q(x, y) \rightarrow X(y)) \rightarrow X(x)).$$

This means, of course, that the two sentences are equivalent. Here's how you might convince yourself of this directly: First suppose the second-order sentence is true in a **structure** M . Since x and X is universally quantified, the remainder must hold for any $x \in W$ and set $X \subseteq W$, e.g, the set $\{z : Rzx\}$ where $R = Q^{\mathfrak{M}}$. So, for any s with $s(x) \in W$ and $s(X) = \{z : Rzx\}$ we have $\mathfrak{M} \models \forall y (Q(x, y) \rightarrow X(y)) \rightarrow X(x)$. But by the way we've picked $s(X)$ that means $\mathfrak{M}, s \models \forall y (Q(x, y) \rightarrow Q(x, y)) \rightarrow Q(x, x)$, which is equivalent to $Q(x, x)$ since the antecedent is valid. Since $s(x)$ is arbitrary, we have $\mathfrak{M} \models \forall x Q(x, x)$.

Now suppose that $\mathfrak{M} \models Q(x, x)$ and show that $\mathfrak{M} \models \forall X \forall x (\forall y (Q(x, y) \rightarrow X(y)) \rightarrow X(x))$. Pick any assignment s , and assume $\mathfrak{M}, s \models \forall y (Q(x, y) \rightarrow X(y))$. Let s' be the y -variant of s with $s'(y) = x$; we have $\mathfrak{M}, s' \models Q(x, y) \rightarrow X(y)$, i.e., $\mathfrak{M}, s \models Q(x, x) \rightarrow X(x)$. Since $\mathfrak{M} \models \forall x Q(x, x)$, the antecedent is true, and we have $\mathfrak{M}, s \models X(x)$, which is what we needed to show.

Since some definable classes of frames are not first-order definable, not every monadic-second order **sentence** of the form φ' is equivalent to a first-order **sentence**. There is no effective method to decide which ones are.

Chapter 3

Axiomatic Derivations

3.1 Introduction

mod:axs:int:
sec

We have a semantics for the basic modal language in terms of modal models, and a notion of a **formula** being valid—true at all worlds in all models—or valid with respect to some class of models or frames—true at all worlds in all models in the class, or based on the frame. Logic usually connects such semantic characterizations of validity with a proof-theoretic notion of **derivability**. The aim is to define a notion of **derivability** in some system such that a **formula** is **derivable** iff it is valid.

The simplest and historically oldest **derivation** systems are so-called Hilbert-type or axiomatic **derivation** systems. Hilbert-type **derivation** systems for many modal logics are relatively easy to construct: they are simple as objects of metatheoretical study (e.g., to prove soundness and completeness). However, they are much harder to use to prove **formulas** in than, say, natural deduction systems.

In Hilbert-type **derivation** systems, a derivation of a **formula** is a sequence of **formulas** leading from certain axioms, via a handful of inference rules, to the **formula** in question. Since we want the **derivation** system to match the semantics, we have to guarantee that the set of **derivable** formulas are true in all models (or true in all models in which all axioms are true). We'll first isolate some properties of modal logics that are necessary for this to work: the “normal” modal logics. For normal modal logics, there are only two inference rules that need to be assumed: modus ponens and necessitation. As axioms we take all (substitution instances) of tautologies, and, depending on the modal logic we deal with, a number of modal axioms. Even if we are just interested in the class of all models, we must also count all substitution instances of **K** and **Dual** as axioms. This alone generates the minimal normal modal logic **K**.

Definition 3.1. The rule of *modus ponens* is the inference schema

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

We say a **formula** ψ follows from **formulas** φ, χ by modus ponens iff $\chi \equiv \varphi \rightarrow \psi$.

Definition 3.2. The rule of *necessitation* is the inference schema

$$\frac{\varphi}{\Box\varphi} \text{ NEC}$$

We say the **formula** ψ follows from the **formulas** φ by necessitation iff $\psi \equiv \Box\varphi$.

Definition 3.3. A *derivation* from a set of axioms Σ is a sequence of **formulas** $\psi_1, \psi_2, \dots, \psi_n$, where each ψ_i is either

1. a substitution instance of a tautology, or
2. a substitution instance of a **formula** in Σ , or
3. follows from two **formulas** ψ_j, ψ_k with $j, k < i$ by modus ponens, or
4. follows from a **formula** ψ_j with $j < i$ by necessitation.

If there is such a **derivation** with $\psi_n \equiv \varphi$, we say that φ is *derivable* from Σ , in symbols $\Sigma \vdash \varphi$.

With this definition, it will turn out that the set of **derivable** formulas forms a normal modal logic, and that any **derivable formula** is true in every model in which every axiom is true. This property of **derivations** is called *soundness*. The converse, *completeness*, is harder to prove.

3.2 Normal Modal Logics

Not every set of modal **formulas** can easily be characterized as those **formulas** derivable from a set of axioms. We want modal logics to be well-behaved. First of all, everything we can **derive** in classical propositional logic should still be **derivable**, of course taking into account that the **formulas** may now contain also \Box and \Diamond . To this end, we require that a modal logic contain all tautological instances and be closed under modus ponens.

mod:prf:nor:
sec

Definition 3.4. A *modal logic* is a set Σ of modal **formulas** which

1. contains all tautologies, and
2. is closed under substitution, i.e., if $\varphi \in \Sigma$, and $\theta_1, \dots, \theta_n$ are **formulas**, then

$$\varphi[\theta_1/p_1, \dots, \theta_n/p_n] \in \Sigma,$$

3. is closed under *modus ponens*, i.e., if φ and $\varphi \rightarrow \psi \in \Sigma$, then $\psi \in \Sigma$.

In order to use the relational semantics for modal logics, we also have to require that all **formulas** valid in all modal models are included. It turns out that this requirement is met as soon as all instances of K and DUAL are **derivable**, and whenever a **formula** φ is **derivable**, so is $\Box\varphi$. A modal logic that satisfies these conditions is called *normal*. (Of course, there are also non-normal modal logics, but the usual relational models are not adequate for them.)

Definition 3.5. A modal logic Σ is *normal* if it contains

$$\Box(p \rightarrow q) \rightarrow (\Box p \rightarrow \Box q), \quad (\text{K})$$

$$\Diamond p \leftrightarrow \neg \Box \neg p \quad (\text{DUAL})$$

and is closed under *necessitation*, i.e., if $\varphi \in \Sigma$, then $\Box \varphi \in \Sigma$.

Observe that while tautological implication is “fine-grained” enough to preserve *truth at a world*, the rule NEC only preserves *truth in a model* (and hence also validity in a frame or in a class of frames).

mod:prf:nor: **Proposition 3.6.** *Every normal modal logic is closed under rule RK,*
prop:rk

$$\frac{\varphi_1 \rightarrow (\varphi_2 \rightarrow \cdots (\varphi_{n-1} \rightarrow \varphi_n) \cdots)}{\Box \varphi_1 \rightarrow (\Box \varphi_2 \rightarrow \cdots (\Box \varphi_{n-1} \rightarrow \Box \varphi_n) \cdots)} \text{RK}$$

Proof. By induction on n : If $n = 1$, then the rule is just NEC, and every normal modal logic is closed under NEC.

Now suppose the result holds for $n - 1$; we show it holds for n .

Assume

$$\varphi_1 \rightarrow (\varphi_2 \rightarrow \cdots (\varphi_{n-1} \rightarrow \varphi_n) \cdots) \in \Sigma$$

By the induction hypothesis, we have

$$\Box \varphi_1 \rightarrow (\Box \varphi_2 \rightarrow \cdots \Box (\varphi_{n-1} \rightarrow \varphi_n) \cdots) \in \Sigma$$

Since Σ is a normal modal logic, it contains all instances of K, in particular

$$\Box (\varphi_{n-1} \rightarrow \varphi_n) \rightarrow (\Box \varphi_{n-1} \rightarrow \Box \varphi_n) \in \Sigma$$

Using modus ponens and suitable tautological instances we get

$$\Box \varphi_1 \rightarrow (\Box \varphi_2 \rightarrow \cdots (\Box \varphi_{n-1} \rightarrow \Box \varphi_n) \cdots) \in \Sigma. \quad \square$$

mod:prf:nor: **Proposition 3.7.** *Every normal modal logic Σ contains $\neg \Diamond \perp$.*
prop:notDiamondBot

Problem 3.1. Prove Proposition 3.7.

Proposition 3.8. *Let $\varphi_1, \dots, \varphi_n$ be formulas. Then there is a smallest modal logic Σ containing all instances of $\varphi_1, \dots, \varphi_n$.*

Proof. Given $\varphi_1, \dots, \varphi_n$, define Σ as the intersection of all normal modal logics containing all instances of $\varphi_1, \dots, \varphi_n$. The intersection is non-empty as $\text{Frm}(\mathcal{L})$, the set of all formulas, is such a modal logic. \square

Definition 3.9. The smallest normal modal logic containing $\varphi_1, \dots, \varphi_n$ is called a *modal system* and denoted by $\mathbf{K}\varphi_1 \dots \varphi_n$. The smallest normal modal logic is denoted by \mathbf{K} .

3.3 Derivations and Modal Systems

We first define what a **derivation** is for normal modal logics. Roughly, a **derivation** is a sequence of **formulas** in which every **element** is either (a substitution instance of) one of a number of *axioms*, or follows from previous **elements** by one of a few inference rules. For normal modal logics, all instances of tautologies, K, and DUAL count as axioms. This results in the modal system **K**, the smallest normal modal logic. We may wish to add additional axioms to obtain other systems, however. The rules are always modus ponens MP and necessitation NEC.

Definition 3.10. Given a modal system $\mathbf{K}\varphi_1 \dots \varphi_n$ and a **formula** ψ we say that ψ is *derivable* in $\mathbf{K}\varphi_1 \dots \varphi_n$, written $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi$, if and only if there are **formulas** χ_1, \dots, χ_k such that $\chi_k = \psi$ and each χ_i is either a tautological instance, or an instance of one of K, DUAL, $\varphi_1, \dots, \varphi_n$, or it follows from previous **formulas** by means of the rules MP or NEC.

The following proposition allows us to show that $\psi \in \Sigma$ by exhibiting a Σ -proof of ψ .

Proposition 3.11. $\mathbf{K}\varphi_1 \dots \varphi_n = \{\psi : \mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi\}$.

Proof. We use induction on the length of **derivations** to show that $\{\psi : \mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi\} \subseteq \mathbf{K}\varphi_1 \dots \varphi_n$.

If the **derivation** of ψ has length 1, it contains a single **formula**. That **formula** cannot follow from previous formulas by MP or NEC, so must be a tautological instance, an instance of K, DUAL, or an instance of one of $\varphi_1, \dots, \varphi_n$. But $\mathbf{K}\varphi_1 \dots \varphi_n$ contains these as well, so $\psi \in \mathbf{K}\varphi_1 \dots \varphi_n$.

If the **derivation** of ψ has length > 1 , then ψ may in addition be obtained by MP or NEC from **formulas** not occurring as the last line in the **derivation**. If ψ follows from χ and $\chi \rightarrow \psi$ (by MP), then χ and $\chi \rightarrow \psi \in \mathbf{K}\varphi_1 \dots \varphi_n$ by induction hypothesis. But every modal logic is closed under modus ponens, so $\psi \in \mathbf{K}\varphi_1 \dots \varphi_n$. If $\psi \equiv \Box\chi$ follows from χ by NEC, then $\chi \in \mathbf{K}\varphi_1 \dots \varphi_n$ by induction hypothesis. But every normal modal logic is closed under NEC, so $\psi \in \mathbf{K}\varphi_1 \dots \varphi_n$.

The converse inclusion follows by showing that $\Sigma = \{\psi : \mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi\}$ is a normal modal logic containing all the instances of $\varphi_1, \dots, \varphi_n$, and the observation that $\mathbf{K}\varphi_1 \dots \varphi_n$ is, by definition, the smallest such logic.

1. Every tautology ψ is a tautological instance, so $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi$, so Σ contains all tautologies.
2. If $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \chi$ and $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \chi \rightarrow \psi$, then $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi$: Combine the **derivation** of χ with that of $\chi \rightarrow \psi$, and add the line ψ . The last line is justified by MP. So Σ is closed under modus ponens.
3. If ψ has a **derivation**, then every substitution instance of ψ also has a derivation: apply the substitution to every **formula** in the **derivation**.

(Exercise: prove by induction on the length of **derivations** that the result is also a correct **derivation**). So Σ is closed under uniform substitution. (We have now established that Σ satisfies all conditions of a modal logic.)

4. We have $\mathbf{K}\varphi_1 \dots \varphi_n \vdash K$, so $K \in \Sigma$.
5. We have $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \text{DUAL}$, so $\text{DUAL} \in \Sigma$.
6. If $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \chi$, the additional line $\Box\chi$ is justified by NEC. Consequently, Σ is closed under NEC. Thus, Σ is normal.

□

3.4 Proofs in K

mod:prf:prk:
sec In order to practice proofs in the smallest modal system, we show the valid **formulas** on the left-hand side of [table 1.1](#) can all be given **K**-proofs.

Proposition 3.12. $\mathbf{K} \vdash \Box\varphi \rightarrow \Box(\psi \rightarrow \varphi)$

Proof.

1. $\varphi \rightarrow (\psi \rightarrow \varphi)$ TAUT
2. $\Box(\varphi \rightarrow (\psi \rightarrow \varphi))$ NEC, 1
3. $\Box(\varphi \rightarrow (\psi \rightarrow \varphi)) \rightarrow (\Box\varphi \rightarrow \Box(\psi \rightarrow \varphi))$ K
4. $\Box\varphi \rightarrow \Box(\psi \rightarrow \varphi)$ MP, 2, 3

□

Proposition 3.13. $\mathbf{K} \vdash \Box(\varphi \wedge \psi) \rightarrow (\Box\varphi \wedge \Box\psi)$

Proof.

1. $(\varphi \wedge \psi) \rightarrow \varphi$ TAUT
2. $\Box((\varphi \wedge \psi) \rightarrow \varphi)$ NEC
3. $\Box((\varphi \wedge \psi) \rightarrow \varphi) \rightarrow (\Box(\varphi \wedge \psi) \rightarrow \Box\varphi)$ K
4. $\Box(\varphi \wedge \psi) \rightarrow \Box\varphi$ MP, 2, 3
5. $(\varphi \wedge \psi) \rightarrow \psi$ TAUT
6. $\Box((\varphi \wedge \psi) \rightarrow \psi)$ NEC
7. $\Box((\varphi \wedge \psi) \rightarrow \psi) \rightarrow (\Box(\varphi \wedge \psi) \rightarrow \Box\psi)$ K
8. $\Box(\varphi \wedge \psi) \rightarrow \Box\psi$ MP, 6, 7
9. $(\Box(\varphi \wedge \psi) \rightarrow \Box\varphi) \rightarrow$
 $((\Box(\varphi \wedge \psi) \rightarrow \Box\varphi) \rightarrow$
 $(\Box(\varphi \wedge \psi) \rightarrow (\Box\varphi \wedge \Box\psi)))$ TAUT
10. $(\Box(\varphi \wedge \psi) \rightarrow \Box\varphi) \rightarrow$
 $(\Box(\varphi \wedge \psi) \rightarrow (\Box\varphi \wedge \Box\psi))$ MP, 4, 9
11. $\Box(\varphi \wedge \psi) \rightarrow (\Box\varphi \wedge \Box\psi)$ MP, 4, 10.

Note that the **formula** on line 9 is an instance of the tautology

$$(p \rightarrow q) \rightarrow ((p \rightarrow r) \rightarrow (p \rightarrow (q \wedge r))).$$

□

Proposition 3.14. $\mathbf{K} \vdash (\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$

Proof.

- | | | |
|-----|---|----------|
| 1. | $\varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi))$ | TAUT |
| 2. | $\Box(\varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi)))$ | NEC, 1 |
| 3. | $\Box(\varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi))) \rightarrow (\Box\varphi \rightarrow \Box(\psi \rightarrow (\varphi \wedge \psi)))$ | K |
| 4. | $\Box\varphi \rightarrow \Box(\psi \rightarrow (\varphi \wedge \psi))$ | MP, 2, 3 |
| 5. | $\Box(\psi \rightarrow (\varphi \wedge \psi)) \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))$ | K |
| 6. | $(\Box\varphi \rightarrow \Box(\psi \rightarrow (\varphi \wedge \psi))) \rightarrow$
$(\Box(\psi \rightarrow (\varphi \wedge \psi)) \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))) \rightarrow$
$(\Box\varphi \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi)))$ | TAUT |
| 7. | $(\Box(\psi \rightarrow (\varphi \wedge \psi)) \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))) \rightarrow$
$(\Box\varphi \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi)))$ | MP, 4, 6 |
| 8. | $\Box\varphi \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))$ | MP, 5, 7 |
| 9. | $(\Box\varphi \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))) \rightarrow$
$((\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi))$ | TAUT |
| 10. | $(\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$ | MP, 8, 9 |

The **formulas** on lines 6 and 9 are instances of the tautologies

$$(p \rightarrow q) \rightarrow ((q \rightarrow r) \rightarrow (p \rightarrow r))$$

$$(p \rightarrow (q \rightarrow r)) \rightarrow ((p \wedge q) \rightarrow r)$$

□

Proposition 3.15. $\mathbf{K} \vdash \neg\Box p \rightarrow \Diamond\neg p$

Proof.

- | | | |
|-----|--|-----------|
| 1. | $\Diamond\neg p \leftrightarrow \neg\Box\neg\neg p$ | DUAL |
| 2. | $(\Diamond\neg p \leftrightarrow \neg\Box\neg\neg p) \rightarrow$
$(\neg\Box\neg\neg p \rightarrow \Diamond\neg p)$ | TAUT |
| 3. | $\neg\Box\neg\neg p \rightarrow \Diamond\neg p$ | MP, 1, 2 |
| 4. | $\neg\neg p \rightarrow p$ | TAUT |
| 5. | $\Box(\neg\neg p \rightarrow p)$ | NEC, 4 |
| 6. | $\Box(\neg\neg p \rightarrow p) \rightarrow (\Box\neg\neg p \rightarrow \Box p)$ | K |
| 7. | $(\Box\neg\neg p \rightarrow \Box p)$ | MP, 5, 6 |
| 8. | $(\Box\neg\neg p \rightarrow \Box p) \rightarrow (\neg\Box p \rightarrow \neg\Box\neg\neg p)$ | TAUT |
| 9. | $\neg\Box p \rightarrow \neg\Box\neg\neg p$ | MP, 7, 8 |
| 10. | $(\neg\Box p \rightarrow \neg\Box\neg\neg p) \rightarrow$
$((\neg\Box\neg\neg p \rightarrow \Diamond\neg p) \rightarrow (\neg\Box p \rightarrow \Diamond\neg p))$ | TAUT |
| 11. | $(\neg\Box\neg\neg p \rightarrow \Diamond\neg p) \rightarrow (\neg\Box p \rightarrow \Diamond\neg p)$ | MP, 9, 10 |
| 12. | $\neg\Box p \rightarrow \Diamond\neg p$ | MP, 3, 11 |

The **formulas** on lines 8 and 10 are instances of the tautologies

$$(p \rightarrow q) \rightarrow (\neg q \rightarrow \neg p)$$

$$(p \rightarrow q) \rightarrow ((q \rightarrow r) \rightarrow (p \rightarrow r)).$$

□

Problem 3.2. Find **derivations** in **K** for the following **formulas**:

1. $\Box\neg p \rightarrow \Box(p \rightarrow q)$
2. $(\Box p \vee \Box q) \rightarrow \Box(p \vee q)$
3. $\Diamond p \rightarrow \Diamond(p \vee q)$

3.5 Derived Rules

mod:prf:der:
sec

Finding and writing **derivations** is obviously difficult, cumbersome, and repetitive. For instance, very often we want to pass from $\varphi \rightarrow \psi$ to $\Box\varphi \rightarrow \Box\psi$, i.e., apply rule RK. That requires an application of NEC, then recording the proper instance of K, then applying MP. Passing from $\varphi \rightarrow \psi$ and $\psi \rightarrow \chi$ to $\varphi \rightarrow \chi$ requires recording the (long) tautological instance

$$(\varphi \rightarrow \psi) \rightarrow ((\psi \rightarrow \chi) \rightarrow (\varphi \rightarrow \chi))$$

and applying MP twice. Often we want to replace a sub-**formula** by a formula we know to be equivalent, e.g., $\Diamond\varphi$ by $\neg\Box\neg\varphi$, or $\neg\neg\varphi$ by φ . So rather than write out the actual **derivation**, it is more convenient to simply record why the intermediate steps are **derivable**. For this purpose, let us collect some facts about **derivability**.

Proposition 3.16. *If $\mathbf{K} \vdash \varphi_1, \dots, \mathbf{K} \vdash \varphi_n$, and ψ follows from $\varphi_1, \dots, \varphi_n$ by propositional logic, then $\mathbf{K} \vdash \psi$.*

Proof. If ψ follows from $\varphi_1, \dots, \varphi_n$ by propositional logic, then

$$\varphi_1 \rightarrow (\varphi_2 \rightarrow \dots (\varphi_n \rightarrow \psi) \dots)$$

is a tautological instance. Applying MP n times gives a **derivation** of ψ . □

We will indicate use of this proposition by PL.

Proposition 3.17. *If $\mathbf{K} \vdash \varphi_1 \rightarrow (\varphi_2 \rightarrow \dots (\varphi_{n-1} \rightarrow \varphi_n) \dots)$ then $\mathbf{K} \vdash \Box\varphi_1 \rightarrow (\Box\varphi_2 \rightarrow \dots (\Box\varphi_{n-1} \rightarrow \Box\varphi_n) \dots)$.*

Proof. By induction on n , just as in the proof of [Proposition 3.6](#). □

We will indicate use of this proposition by RK. Let's illustrate how these results help establishing **derivability** results more easily.

Proposition 3.18. $\mathbf{K} \vdash (\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$

Proof.

1. $\mathbf{K} \vdash \varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi))$ TAUT
2. $\mathbf{K} \vdash \Box\varphi \rightarrow (\Box\psi \rightarrow \Box(\varphi \wedge \psi))$ RK, 1
3. $\mathbf{K} \vdash (\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$ PL, 2

□

Proposition 3.19. *If $\mathbf{K} \vdash \varphi \leftrightarrow \psi$ and $\mathbf{K} \vdash \chi[\varphi/p]$ then $\mathbf{K} \vdash \chi[B/p]$*

*mod:prf:der:
prop:rewriting*

Proof. Exercise.

□

Problem 3.3. Prove [Proposition 3.19](#) by proving, by induction on the complexity of χ , that if $\mathbf{K} \vdash \varphi \leftrightarrow \psi$ then $\mathbf{K} \vdash \chi[\varphi/p] \leftrightarrow \chi[\psi/p]$.

This proposition comes in handy especially when we want to convert \Diamond into \Box (or vice versa), or remove double negations inside a [formula](#). For instance:

Proposition 3.20. $\mathbf{K} \vdash \neg\Box p \rightarrow \Diamond\neg p$

Proof.

1. $\mathbf{K} \vdash \Diamond\neg p \leftrightarrow \neg\Box\neg\neg p$ DUAL
2. $\mathbf{K} \vdash \neg\Box\neg\neg p \rightarrow \Diamond\neg p$ PL, 1
3. $\mathbf{K} \vdash \neg\Box p \rightarrow \Diamond\neg p$ re-write p for $\neg\neg p$

□

The following proposition justifies that we can establish [derivability](#) results schematically. E.g., the previous proposition does not just establish that $\mathbf{K} \vdash \neg\Box p \rightarrow \Diamond\neg p$, but $\mathbf{K} \vdash \neg\Box\varphi \rightarrow \Diamond\neg\varphi$ for arbitrary φ .

Proposition 3.21. *If φ is a substitution instance of ψ and $\mathbf{K} \vdash \psi$, then $\mathbf{K} \vdash \varphi$.*

Proof. It is tedious but routine to verify (by induction on the length of the [derivation](#) of ψ) that applying a substitution to an entire [derivation](#) also results in a correct [derivation](#). Specifically, substitution instances of tautological instances are themselves tautological instances, substitution instances of instances of DUAL and K are themselves instances of DUAL and K, and applications of MP and NEC remain correct when substituting [formulas](#) for [propositional variables](#) in both premise(s) and conclusion. □

3.6 More Proofs in \mathbf{K}

Let's see some more examples of [derivability](#) in \mathbf{K} , now using the simplified method introduced in [section 3.5](#).

*mod:prf:mpr:
sec*

Proposition 3.22. $\mathbf{K} \vdash \Box(\varphi \rightarrow \psi) \rightarrow (\Diamond\varphi \rightarrow \Diamond\psi)$

Proof.

1. $\mathbf{K} \vdash (\varphi \rightarrow \psi) \rightarrow (\neg\psi \rightarrow \neg\varphi)$ PL
2. $\mathbf{K} \vdash \Box(\varphi \rightarrow \psi) \rightarrow (\Box\neg\psi \rightarrow \Box\neg\varphi)$ RK, 1
3. $\mathbf{K} \vdash (\Box\neg\psi \rightarrow \Box\neg\varphi) \rightarrow (\neg\Box\neg\varphi \rightarrow \neg\Box\neg\psi)$ TAUT
4. $\mathbf{K} \vdash (\Box\neg\psi \rightarrow \Box\neg\varphi) \rightarrow (\neg\Box\neg\varphi \rightarrow \neg\Box\neg\psi)$ PL, 2, 3
5. $\mathbf{K} \vdash \Box(\varphi \rightarrow \psi) \rightarrow (\Diamond\varphi \rightarrow \Diamond\psi)$ re-writing \Diamond for $\neg\Box\neg$.

□

Proposition 3.23. $\mathbf{K} \vdash \Box\varphi \rightarrow (\Diamond(\varphi \rightarrow \psi) \rightarrow \Diamond\psi)$

Proof.

1. $\mathbf{K} \vdash \varphi \rightarrow (\neg\psi \rightarrow \neg(\varphi \rightarrow \psi))$ TAUT
2. $\mathbf{K} \vdash \Box\varphi \rightarrow (\Box\neg\psi \rightarrow \Box\neg(\varphi \rightarrow \psi))$ RK, 1
3. $\mathbf{K} \vdash \Box\varphi \rightarrow (\neg\Box\neg(\varphi \rightarrow \psi) \rightarrow \neg\Box\neg\psi)$ PL, 2
4. $\mathbf{K} \vdash \Box\varphi \rightarrow (\Diamond(\varphi \rightarrow \psi) \rightarrow \Diamond\psi)$ re-writing \Diamond for $\neg\Box\neg$.

□

Proposition 3.24. $\mathbf{K} \vdash (\Diamond\varphi \vee \Diamond\psi) \rightarrow \Diamond(\varphi \vee \psi)$

Proof.

1. $\mathbf{K} \vdash \neg(\varphi \vee \psi) \rightarrow \neg\varphi$ TAUT
2. $\mathbf{K} \vdash \Box\neg(\varphi \vee \psi) \rightarrow \Box\neg\varphi$ RK, 1
3. $\mathbf{K} \vdash \neg\Box\neg\varphi \rightarrow \neg\Box\neg(\varphi \vee \psi)$ PL, 2
4. $\mathbf{K} \vdash \Diamond\varphi \rightarrow \Diamond(\varphi \vee \psi)$ re-writing
5. $\mathbf{K} \vdash \Diamond\psi \rightarrow \Diamond(\varphi \vee \psi)$ similarly
6. $\mathbf{K} \vdash (\Diamond\varphi \vee \Diamond\psi) \rightarrow \Diamond(\varphi \vee \psi)$ PL, 4, 5.

□

Proposition 3.25. $\mathbf{K} \vdash \Diamond(\varphi \vee \psi) \rightarrow (\Diamond\varphi \vee \Diamond\psi)$

Proof.

1. $\mathbf{K} \vdash \neg\varphi \rightarrow (\neg\psi \rightarrow \neg(\varphi \vee \psi))$ TAUT
2. $\mathbf{K} \vdash \Box\neg\varphi \rightarrow (\Box\neg\psi \rightarrow \Box\neg(\varphi \vee \psi))$ RK
3. $\mathbf{K} \vdash \Box\neg\varphi \rightarrow (\neg\Box\neg(\varphi \vee \psi) \rightarrow \neg\Box\neg\psi)$ PL, 2
4. $\mathbf{K} \vdash \neg\Box\neg(\varphi \vee \psi) \rightarrow (\Box\neg\varphi \rightarrow \neg\Box\neg\psi)$ PL, 3
5. $\mathbf{K} \vdash \neg\Box\neg(\varphi \vee \psi) \rightarrow (\neg\neg\Box\neg\psi \rightarrow \neg\Box\neg\varphi)$ PL, 4
6. $\mathbf{K} \vdash \Diamond(\varphi \vee \psi) \rightarrow (\neg\Diamond\psi \rightarrow \Diamond\varphi)$ re-writing \Diamond for $\neg\Box\neg$
7. $\mathbf{K} \vdash \Diamond(\varphi \vee \psi) \rightarrow (\Diamond\psi \vee \Diamond\varphi)$ PL, 6.

□

Problem 3.4. Show that the following **derivability** claims hold:

1. $\mathbf{K} \vdash \Diamond\neg\perp \rightarrow (\Box\varphi \rightarrow \Diamond\varphi)$;
2. $\mathbf{K} \vdash \Box(\varphi \vee \psi) \rightarrow (\Diamond\varphi \vee \Box\psi)$;
3. $\mathbf{K} \vdash (\Diamond\varphi \rightarrow \Box\psi) \rightarrow \Box(\varphi \rightarrow \psi)$.

3.7 Dual Formulas

Definition 3.26. Each of the formulas T, B, 4, and 5 has a *dual*, denoted by a subscripted diamond, as follows:

mod:prf:dua:
sec
mod:prf:dua:
def:duals

$$\begin{aligned} p &\rightarrow \Diamond p && (T_{\Diamond}) \\ \Diamond \Box p &\rightarrow p && (B_{\Diamond}) \\ \Diamond \Diamond p &\rightarrow \Diamond p && (4_{\Diamond}) \\ \Diamond \Box p &\rightarrow \Box p && (5_{\Diamond}) \end{aligned}$$

Each of the above dual formulas is obtained from the corresponding formula by substituting $\neg p$ for p , contraposing, replacing $\neg \Box \neg$ by \Diamond , and replacing $\neg \Diamond \neg$ by \Box . D, i.e., $\Box \varphi \rightarrow \Diamond \varphi$ is its own dual in that sense.

Problem 3.5. Show that for each formula φ in Definition 3.26: $\mathbf{K} \vdash \varphi \leftrightarrow \varphi_{\Diamond}$.

3.8 Proofs in Modal Systems

We now come to proofs in systems of modal logic other than **K**.

Proposition 3.27. *The following provability results obtain:*

mod:prf:prs:
sec
mod:prf:prs:
prop:S5facts

1. **KT5** \vdash B;
2. **KT5** \vdash 4;
3. **KDB4** \vdash T;
4. **KB4** \vdash 5;
5. **KB5** \vdash 4;
6. **KT** \vdash D.

mod:prf:prs:
prop:S5facts-KT-D

Proof. We exhibit proofs for each.

1. **KT5** \vdash B:

1. **KT5** $\vdash \Diamond \varphi \rightarrow \Box \Diamond \varphi$ 5
2. **KT5** $\vdash \varphi \rightarrow \Diamond \varphi$ T_{\Diamond}
3. **KT5** $\vdash \varphi \rightarrow \Box \Diamond \varphi$ PL.

2. **KT5** \vdash 4:

1. **KT5** $\vdash \Diamond \Box \varphi \rightarrow \Box \Diamond \Box \varphi$ 5 with $\Box \varphi$ for p
2. **KT5** $\vdash \Box \varphi \rightarrow \Diamond \Box \varphi$ T_{\Diamond} with $\Box \varphi$ for p
3. **KT5** $\vdash \Box \varphi \rightarrow \Box \Diamond \Box \varphi$ PL, 1, 2
4. **KT5** $\vdash \Diamond \Box \varphi \rightarrow \Box \varphi$ 5_{\Diamond}
5. **KT5** $\vdash \Box \Diamond \Box \varphi \rightarrow \Box \Box \varphi$ RK, 4
6. **KT5** $\vdash \Box \varphi \rightarrow \Box \Box \varphi$ PL, 3, 5.

3. **KDB4** ⊢ T:

1. **KDB4** ⊢ $\Diamond\Box\varphi \rightarrow \varphi$ B_\Diamond
2. **KDB4** ⊢ $\Box\Box\varphi \rightarrow \Diamond\Box\varphi$ D with $\Box\varphi$ for p
3. **KDB4** ⊢ $\Box\Box\varphi \rightarrow \varphi$ PL1, 2
4. **KDB4** ⊢ $\Box\varphi \rightarrow \Box\Box\varphi$ 4
5. **KDB4** ⊢ $\Box\varphi \rightarrow \varphi$ PL, 1, 4.

4. **KB4** ⊢ 5:

1. **KB4** ⊢ $\Diamond\varphi \rightarrow \Box\Diamond\Diamond\varphi$ B with $\Diamond\varphi$ for p
2. **KB4** ⊢ $\Diamond\Diamond\varphi \rightarrow \Diamond\varphi$ 4_\Diamond
3. **KB4** ⊢ $\Box\Diamond\Diamond\varphi \rightarrow \Box\Diamond\varphi$ RK, 2
4. **KB4** ⊢ $\Diamond\varphi \rightarrow \Box\Diamond\varphi$ PL, 1, 3.

5. **KB5** ⊢ 4:

1. **KB5** ⊢ $\Box\varphi \rightarrow \Box\Diamond\Box\varphi$ B with $\Box\varphi$ for p
2. **KB5** ⊢ $\Diamond\Box\varphi \rightarrow \Box\varphi$ 5_\Diamond
3. **KB5** ⊢ $\Box\Diamond\Box\varphi \rightarrow \Box\Box\varphi$ RK, 2
4. **KB5** ⊢ $\Box\varphi \rightarrow \Box\Box\varphi$ PL, 1, 3.

6. **KT** ⊢ D:

1. **KT** ⊢ $\Box\varphi \rightarrow \varphi$ T
2. **KT** ⊢ $\varphi \rightarrow \Diamond\varphi$ T_\Diamond
3. **KT** ⊢ $\Box\varphi \rightarrow \Diamond\varphi$ PL, 1, 2

□

Definition 3.28. Following tradition, we define **S4** to be the system **KT4**, and **S5** the system **KTB4**.

The following proposition shows that the classical system **S5** has several equivalent axiomatizations. This should not surprise, as the various combinations of axioms all characterize equivalence relations (see [Proposition 2.12](#)).

mod:prf:prs: **Proposition 3.29.** **KTB4 = KT5 = KDB4 = KDB5.**
prop:S5

Proof. Exercise.

□

Problem 3.6. Prove [Proposition 3.29](#).

3.9 Soundness

A derivation system is called sound if everything that can be derived is valid. When considering modal systems, i.e., derivations where in addition to \mathbf{K} we can use instances of some formulas $\varphi_1, \dots, \varphi_n$, we want every derivable formula to be true in any model in which $\varphi_1, \dots, \varphi_n$ are true. mod:prf:snd:
sec

Theorem 3.30 (Soundness Theorem). *If every instance of $\varphi_1, \dots, \varphi_n$ is valid in the classes of models $\mathcal{C}_1, \dots, \mathcal{C}_n$, respectively, then $\mathbf{K}\varphi_1 \dots \varphi_n \vdash \psi$ implies that ψ is valid in the class of models $\mathcal{C}_1 \cap \dots \cap \mathcal{C}_n$.* mod:prf:snd:
thm:soundness

Proof. By induction on length of proofs. For brevity, put $\mathcal{C} = \mathcal{C}_1 \cap \dots \cap \mathcal{C}_n$.

1. Induction Basis: If ψ has a proof of length 1, then it is either a tautological instance, an instance of \mathbf{K} , or of DUAL, or an instance of one of $\varphi_1, \dots, \varphi_n$. In the first case, ψ is valid in \mathcal{C} , since tautological instances are valid in any class of models, by Proposition 1.15. Similarly in the second case, by Proposition 1.18 and Proposition 1.19. Finally in the third case, since ψ is valid in \mathcal{C}_i and $\mathcal{C} \subseteq \mathcal{C}_i$, we have that ψ is valid in \mathcal{C} as well.
2. Inductive step: Suppose ψ has a proof of length $k > 1$. If ψ is a tautological instance or an instance of one of $\varphi_1, \dots, \varphi_n$, we proceed as in the previous step. So suppose ψ is obtained by MP from previous formulas $\chi \rightarrow \psi$ and χ . Then $\chi \rightarrow \psi$ and χ have proofs of length $< k$, and by inductive hypothesis they are valid in \mathcal{C} . By Proposition 1.20, ψ is valid in \mathcal{C} as well. Finally suppose ψ is obtained by NEC from χ (so that $\psi = \Box\chi$). By inductive hypothesis, χ is valid in \mathcal{C} , and by Proposition 1.12 so is ψ . \square

3.10 Showing Systems are Distinct

In section 3.8 we saw how to prove that two systems of modal logic are in fact the same system. Theorem 3.30 allows us to show that two modal systems Σ and Σ' are distinct, by finding a formula φ such that $\Sigma' \vdash \varphi$ that fails in a model of Σ . mod:prf:dis:
sec

Proposition 3.31. $\mathbf{KD} \subsetneq \mathbf{KT}$

Proof. This is the syntactic counterpart to the semantic fact that all reflexive relations are serial. To show $\mathbf{KD} \subseteq \mathbf{KT}$ we need to see that $\mathbf{KD} \vdash \psi$ implies $\mathbf{KT} \vdash \psi$, which follows from $\mathbf{KT} \vdash \mathbf{D}$, as shown in Proposition 3.27(6). To show that the inclusion is proper, by Soundness (Theorem 3.30), it suffices to exhibit a model of \mathbf{KD} where T, i.e., $\Box p \rightarrow p$, fails (an easy task left as an exercise), for then by Soundness $\mathbf{KD} \not\vdash \Box p \rightarrow p$. \square

Proposition 3.32. $\mathbf{KB} \neq \mathbf{K4}$.

Proof. We construct a symmetric model where some instance of 4 fails; since obviously the instance is **derivable** for **K4** but not in **KB**, it will follow $\mathbf{K4} \not\subseteq \mathbf{KB}$. Consider the symmetric model \mathfrak{M} of Figure 3.1. Since the model is symmetric, K and B are true in \mathfrak{M} (by Proposition 1.18 and Theorem 2.1, respectively). However, $\mathfrak{M}, w_1 \not\models \Box p \rightarrow \Box \Box p$. \square

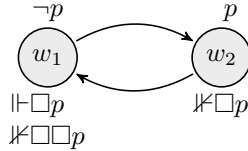


Figure 3.1: A symmetric model falsifying an instance of 4.

mod:prf:dis:
fig:3.1
thm:KTBnot45

Theorem 3.33. $\mathbf{KTB} \not\vdash 4$ and $\mathbf{KTB} \not\vdash 5$.

Proof. By Theorem 2.1 we know that all instances of T and B are true in every reflexive symmetric model (respectively). So by soundness, it suffices to find a reflexive symmetric model containing a world at which some instance of 4 fails, and similarly for 5. We use the same model for both claims. Consider the symmetric, reflexive model in Figure 3.2. Then $\mathfrak{M}, w_1 \not\models \Box p \rightarrow \Box \Box p$, so 4 fails at w_1 . Similarly, $\mathfrak{M}, w_2 \not\models \Diamond \neg p \rightarrow \Box \Diamond \neg p$, so the instance of 5 with $\varphi = \neg p$ fails at w_2 . \square

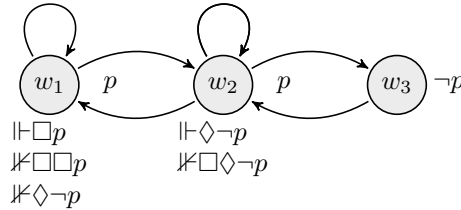


Figure 3.2: The model for Theorem 3.33.

mod:prf:dis:
fig:3.2
thm:KD5not4

Theorem 3.34. $\mathbf{KD5} \neq \mathbf{KT4} = \mathbf{S4}$.

Proof. By Theorem 2.1 we know that all instances of D and 5 are true in all serial euclidean models. So it suffices to find a serial euclidean model containing a world at which some instance of 4 fails. Consider the model of Figure 3.3, and notice that $\mathfrak{M}, w_1 \not\models \Box p \rightarrow \Box \Box p$. \square

Problem 3.7. Give an alternative proof of Theorem 3.34 using a model with 3 worlds.

Problem 3.8. Provide a single reflexive transitive model showing that both $\mathbf{KT4} \not\vdash B$ and $\mathbf{KT4} \not\vdash 5$.

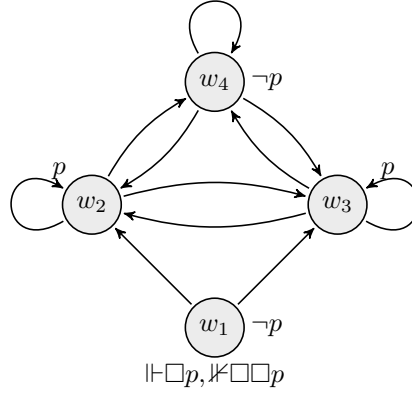


Figure 3.3: The model for Theorem 3.34.

3.11 Derivability from a Set of Formulas

mod:prf:dis:
fig:KD5not4

In section 3.8 we defined a notion of provability of a formula in a system Σ . We now extend this notion to provability in Σ from formulas in a set Γ .

mod:prf:prg:
sec

Definition 3.35. A formula φ is **derivable** in a system Σ from a set of formulas Γ , written $\Gamma \vdash_{\Sigma} \varphi$ if and only if there are $\psi_1, \dots, \psi_n \in \Gamma$ such that $\Sigma \vdash \psi_1 \rightarrow (\psi_2 \rightarrow \dots (\psi_n \rightarrow \varphi) \dots)$.

mod:prf:prg:
defn:Gammaproves

3.12 Properties of Derivability

mod:prf:prp:
sec

Proposition 3.36. Let Σ be a modal system and Γ a set of modal formulas. The following properties hold:

mod:prf:prp:
prop:derivabilityfacts

1. Monotony: If $\Gamma \vdash_{\Sigma} \varphi$ and $\Gamma \subseteq \Delta$ then $\Delta \vdash_{\Sigma} \varphi$;
2. Reflexivity: If $\varphi \in \Gamma$ then $\Gamma \vdash_{\Sigma} \varphi$;
3. Cut: If $\Gamma \vdash_{\Sigma} \varphi$ and $\Delta \cup \{\varphi\} \vdash_{\Sigma} \psi$ then $\Gamma \cup \Delta \vdash_{\Sigma} \psi$;
4. Deduction theorem: $\Gamma \cup \{\psi\} \vdash_{\Sigma} \varphi$ if and only if $\Gamma \vdash_{\Sigma} \psi \rightarrow \varphi$;
5. Rule T: If $\Gamma \vdash_{\Sigma} \varphi_1$ and \dots and $\Gamma \vdash_{\Sigma} \varphi_n$ and $\varphi_1 \rightarrow (\varphi_2 \rightarrow \dots (\varphi_n \rightarrow \psi) \dots)$ is a tautological instance, then $\Gamma \vdash_{\Sigma} \psi$.

mod:prf:prp:
prop:derivabilityfacts-ruleT

The proof is an easy exercise. Part (5) of Proposition 3.36 gives us that, for instance, if $\Gamma \vdash_{\Sigma} \varphi \vee \psi$ and $\Gamma \vdash_{\Sigma} \neg\varphi$, then $\Gamma \vdash_{\Sigma} \psi$. Also, in what follows, we write $\Gamma, \varphi \vdash_{\Sigma} \psi$ instead of $\Gamma \cup \{\varphi\} \vdash_{\Sigma} \psi$.

Definition 3.37. A set Γ is **deductively closed** relatively to a system Σ if and only if $\Gamma \vdash_{\Sigma} \varphi$ implies $\varphi \in \Gamma$.

3.13 Consistency

mod:prf:con:
sec

Consistency is an important property of sets of **formulas**. A set of **formulas** is inconsistent if a contradiction, such as \perp , is **derivable** from it; and otherwise consistent. If a set is inconsistent, its **formulas** cannot all be true in a model at a world. For the completeness theorem we prove the converse: every consistent set is true at a world in a model, namely in the “canonical model.”

Definition 3.38. A set Γ is *consistent* relatively to a system Σ or, as we will say, Σ -consistent, if and only if $\Gamma \not\vdash_{\Sigma} \perp$.

So for instance, the set $\{\Box(p \rightarrow q), \Box p, \neg \Box q\}$ is consistent relative to propositional logic, but not **K**-consistent. Similarly, the set $\{\Diamond p, \Box \Diamond p \rightarrow q, \neg q\}$ is not **K5**-consistent.

mod:prf:con:
prop:consistencyfacts

Proposition 3.39. *Let Γ be a set of **formulas**. Then:*

1. *A set Γ is Σ -consistent if and only if there is some **formula** φ such that $\Gamma \not\vdash_{\Sigma} \varphi$.*
2. *$\Gamma \vdash_{\Sigma} \varphi$ if and only if $\Gamma \cup \{\neg \varphi\}$ is not Σ -consistent.*
3. *If Γ is Σ -consistent, then for any **formula** φ , either $\Gamma \cup \{\varphi\}$ is Σ -consistent or $\Gamma \cup \{\neg \varphi\}$ is Σ -consistent.*

mod:prf:con:
prop:consistencyfacts-b
mod:prf:con:
prop:consistencyfacts-c

Proof. These facts follow easily using classical propositional logic. We give the argument for (3). Proceed contrapositively and suppose neither $\Gamma \cup \{\varphi\}$ nor $\Gamma \cup \{\neg \varphi\}$ is Σ -consistent. Then by (2), both $\Gamma, \varphi \vdash_{\Sigma} \perp$ and $\Gamma, \neg \varphi \vdash_{\Sigma} \perp$. By the deduction theorem $\Gamma \vdash_{\Sigma} \varphi \rightarrow \perp$ and $\Gamma \vdash_{\Sigma} \neg \varphi \rightarrow \perp$. But $(\varphi \rightarrow \perp) \rightarrow ((\neg \varphi \rightarrow \perp) \rightarrow \perp)$ is a tautological instance, hence by [Proposition 3.36\(5\)](#), $\Gamma \vdash_{\Sigma} \perp$. \square

Chapter 4

Completeness and Canonical Models

4.1 Introduction

If Σ is a modal system, then the soundness theorem establishes that if $\Sigma \vdash \varphi$, then φ is valid in any class \mathcal{C} of models in which all instances of all **formulas** in Σ are valid. In particular that means that if $\mathbf{K} \vdash \varphi$ then φ is true in all models; if $\mathbf{KT} \vdash \varphi$ then φ is true in all reflexive models; if $\mathbf{KD} \vdash \varphi$ then φ is true in all serial models, etc. mod:com:int:
sec

Completeness is the converse of soundness: that \mathbf{K} is complete means that if a **formula** φ is valid, $\vdash \varphi$, for instance. Proving completeness is a lot harder to do than proving soundness. It is useful, first, to consider the contrapositive: \mathbf{K} is complete iff whenever $\not\vdash \varphi$, there is a countermodel, i.e., a model \mathfrak{M} such that $\mathfrak{M} \not\models \varphi$. Equivalently (negating φ), we could prove that whenever $\not\vdash \neg\varphi$, there is a model of φ . In the construction of such a model, we can use information contained in φ . When we find models for specific **formulas** we often do the same: E.g., if we want to find a countermodel to $p \rightarrow \Box q$, we know that it has to contain a world where p is true and $\Box q$ is false. And a world where $\Box q$ is false means there has to be a world accessible from it where q is false. And that's all we need to know: which worlds make the **propositional variables** true, and which worlds are accessible from which worlds.

In the case of proving completeness, however, we don't have a specific **formula** φ for which we are constructing a model. We want to establish that a model exists for every φ such that $\not\vdash_{\Sigma} \neg\varphi$. This is a minimal requirement, since if $\vdash_{\Sigma} \neg\varphi$, by soundness, there is no model for φ (in which Σ is true). Now note that $\not\vdash_{\Sigma} \neg\varphi$ iff φ is Σ -consistent. (Recall that $\Sigma \not\vdash_{\Sigma} \neg\varphi$ and $\varphi \not\vdash_{\Sigma} \perp$ are equivalent.) So our task is to construct a model for every Σ -consistent **formula**.

The trick we'll use is to find a Σ -consistent set of **formulas** that contains φ , but also other formulas which tell us what the world that makes φ true has to look like. Such sets are *complete* Σ -consistent sets. It's not enough to construct a model with a single world to make φ true, it will have to contain

multiple worlds and an accessibility relation. The complete Σ -consistent set containing φ will also contain other **formulas** of the form $\Box\psi$ and $\Diamond\chi$. In all accessible worlds, ψ has to be true; in at least one, χ has to be true. In order to accomplish this, we'll simply take *all* possible complete Σ -consistent sets as the basis for the set of worlds. A tricky part will be to figure out when a complete Σ -consistent set should count as being accessible from another in our model.

We'll show that in the model so defined, φ is true at a world—which is also a complete Σ -consistent set—iff φ is **an element** of that set. If φ is Σ -consistent, it will be **an element** of at least one complete Σ -consistent set (a fact we'll prove), and so there will be a world where φ is true. So we will have a single model where every Σ -consistent **formula** φ is true at some world. This single model is the *canonical* model for Σ .

4.2 Complete Σ -Consistent Sets

mod:com:ccs:
sec

Suppose Σ is a set of modal **formulas**—think of them as the axioms or defining principles of a normal modal logic. A set Γ is Σ -consistent iff $\Gamma \not\vdash_{\Sigma} \perp$, i.e., if there is no **derivation** of $\varphi_1 \rightarrow (\varphi_2 \rightarrow \dots (\varphi_n \rightarrow \perp) \dots)$ from Σ , where each $\varphi_i \in \Gamma$. We will construct a “canonical” model in which each world is taken to be a special kind of Σ -consistent set: one which is not just Σ -consistent, but maximally so, in the sense that it settles the truth value of every modal **formula**: for every φ , either $\varphi \in \Gamma$ or $\neg\varphi \in \Gamma$:

Definition 4.1. A set Γ is *complete Σ -consistent* if and only if it is Σ -consistent and for every φ , either $\varphi \in \Gamma$ or $\neg\varphi \in \Gamma$.

Complete Σ -consistent sets Γ have a number of useful properties. For one, they are deductively closed, i.e., if $\Gamma \vdash_{\Sigma} \varphi$ then $\varphi \in \Gamma$. This means in particular that every instance of a **formula** $\varphi \in \Sigma$ is also $\in \Gamma$. Moreover, membership in Γ mirrors the truth conditions for the propositional connectives. This will be important when we define the “canonical model.”

mod:com:ccs:
prop:ccs-properties

Proposition 4.2. *Suppose Γ is complete Σ -consistent. Then:*

mod:com:ccs:
prop:ccs-closed

1. Γ is deductively closed in Σ .

mod:com:ccs:
prop:ccs-sigma

2. $\Sigma \subseteq \Gamma$.

mod:com:ccs:
prop:ccs-lfalse

3. $\perp \notin \Gamma$

mod:com:ccs:
prop:ccs-ltrue

4. $\top \in \Gamma$

mod:com:ccs:
prop:ccs-lnot

5. $\neg\varphi \in \Gamma$ if and only if $\varphi \notin \Gamma$.

mod:com:ccs:
prop:ccs-land

6. $\varphi \wedge \psi \in \Gamma$ iff $\varphi \in \Gamma$ and $\psi \in \Gamma$

mod:com:ccs:
prop:ccs-lor

7. $\varphi \vee \psi \in \Gamma$ iff $\varphi \in \Gamma$ or $\psi \in \Gamma$

8. $\varphi \rightarrow \psi \in \Gamma$ iff $\varphi \notin \Gamma$ or $\psi \in \Gamma$

*mod:com:ccs:
prop:ccs-lif*

9. $\varphi \leftrightarrow \psi \in \Gamma$ iff either $\varphi \in \Gamma$ and $\psi \in \Gamma$, or $\varphi \notin \Gamma$ and $\psi \notin \Gamma$

*mod:com:ccs:
prop:ccs-liff*

Proof. 1. Suppose $\Gamma \vdash_{\Sigma} \varphi$ but $\varphi \notin \Gamma$. Then since Γ is complete Σ -consistent, $\neg\varphi \in \Gamma$. This would make Γ inconsistent, since $\varphi, \neg\varphi \vdash_{\Sigma} \perp$.

2. If $\varphi \in \Sigma$ then $\Gamma \vdash_{\Sigma} \varphi$, and $\varphi \in \Gamma$ by deductive closure, i.e., case (1).

3. If $\perp \in \Gamma$, then $\Gamma \vdash_{\Sigma} \perp$, so Γ would be Σ -inconsistent.

4. $\Gamma \vdash_{\Sigma} \top$, so $\top \in \Gamma$ by deductive closure, i.e., case (1).

5. If $\neg\varphi \in \Gamma$, then by consistency $\varphi \notin \Gamma$; and if $\varphi \notin \Gamma$ then $\varphi \in \Gamma$ since Γ is complete Σ -consistent.

6. Suppose $\varphi \wedge \psi \in \Gamma$. Since $(\varphi \wedge \psi) \rightarrow \varphi$ is a tautological instance, $\varphi \in \Gamma$ by deductive closure, i.e., case (1). Similarly for $\psi \in \Gamma$. On the other hand, suppose both $\varphi \in \Gamma$ and $\psi \in \Gamma$. Then deductive closure implies $(\varphi \wedge \psi) \in \Gamma$, since $\varphi \rightarrow (\psi \rightarrow (\varphi \wedge \psi))$ is a tautological instance.

7. Suppose $\varphi \vee \psi \in \Gamma$, and $\varphi \notin \Gamma$ and $\psi \notin \Gamma$. Since Γ is complete Σ -consistent, $\neg\varphi \in \Gamma$ and $\neg\psi \in \Gamma$. Then $\neg(\varphi \vee \psi) \in \Gamma$ since $\neg\varphi \rightarrow (\neg\psi \rightarrow \neg(\varphi \vee \psi))$ is a tautological instance. This would mean that Γ is Σ -inconsistent, a contradiction.

8. Suppose $\varphi \rightarrow \psi \in \Gamma$ and $\varphi \in \Gamma$; then $\Gamma \vdash_{\Sigma} \psi$, whence $\psi \in \Gamma$ by deductive closure. Conversely, if $\varphi \rightarrow \psi \notin \Gamma$ then since Γ is complete Σ -consistent, $\neg(\varphi \rightarrow \psi) \in \Gamma$. Since $\neg(\varphi \rightarrow \psi) \rightarrow \varphi$ is a tautological instance, $\varphi \in \Gamma$ by deductive closure. Since $\neg(\varphi \rightarrow \psi) \rightarrow \neg\psi$ is a tautological instance, $\neg\psi \in \Gamma$. Then $\psi \notin \Gamma$ since Γ is Σ -consistent.

9. Suppose $\varphi \leftrightarrow \psi \in \Gamma$. If $\varphi \in \Gamma$, then $\psi \in \Gamma$, since $(\varphi \leftrightarrow \psi) \rightarrow (\varphi \rightarrow \psi)$ is a tautological instance. Similarly, if $\psi \in \Gamma$, then $\varphi \in \Gamma$. So either both $\varphi \in \Gamma$ and $\psi \in \Gamma$, or neither $\varphi \in \Gamma$ nor $\psi \in \Gamma$.

Conversely, suppose $\varphi \rightarrow \psi \notin \Gamma$. Since Γ is complete Σ -consistent, $\neg(\varphi \leftrightarrow \psi) \in \Gamma$. Since $\neg(\varphi \leftrightarrow \psi) \rightarrow (\varphi \rightarrow \neg\psi)$ is a tautological instance, if $\varphi \in \Gamma$ then $\neg\psi \in \Gamma$, and since Γ is Σ -consistent, $\psi \notin \Gamma$. Similarly, if $\psi \in \Gamma$ then $\varphi \notin \Gamma$. So neither $\varphi \in \Gamma$ and $\psi \in \Gamma$, nor $\varphi \notin \Gamma$ and $\psi \notin \Gamma$. □

Problem 4.1. Complete the proof of [Proposition 4.2](#).

4.3 Lindenbaum's Lemma

Lindenbaum's Lemma establishes that every Σ -consistent set of **formulas** is contained in at least one *complete* Σ -consistent set. Our construction of the canonical model will show that for each complete Σ -consistent set Δ , there is a world in the canonical model where all and only the **formulas** in Δ are true. So

*mod:com:lin:
sec*

Lindenbaum's Lemma guarantees that every Σ -consistent set is true at some world in the canonical model.

[mod:com:lin:](#)
[thm:lindenbaum](#)

Theorem 4.3 (Lindenbaum's Lemma). *If Γ is Σ -consistent then there is a complete Σ -consistent set Δ extending Γ .*

Proof. Let $\varphi_0, \varphi_1, \dots$ be an exhaustive listing of all formulas of the language (repetitions are allowed). For instance, start by listing ρ_0 , and at each stage $n \geq 1$ list the finitely many formulas of length n using only variables among ρ_0, \dots, ρ_n . We define sets of formulas Δ_n by induction on n , and we then set $\Delta = \bigcup_n \Delta_n$. We first put $\Delta_0 = \Gamma$. Supposing that Δ_n has been defined, we define Δ_{n+1} by:

$$\Delta_{n+1} = \begin{cases} \Delta_n \cup \{\varphi_n\}, & \text{if } \Delta_n \cup \{\varphi_n\} \text{ is consistent;} \\ \Delta_n \cup \{\neg\varphi_n\}, & \text{otherwise.} \end{cases}$$

If we now let $\Delta = \bigcup_{n=0}^{\infty} \Delta_n$.

We have to show that this definition actually yields a set Δ with the required properties, i.e., $\Gamma \subseteq \Delta$ and Δ is complete Σ -consistent.

It's obvious that $\Gamma \subseteq \Delta$, since $\Delta_0 \subseteq \Delta$ by construction, and $\Delta_0 = \Gamma$. In fact, $\Delta_n \subseteq \Delta$ for all n , since Δ is the union of all Δ_n . (Since in each step of the construction, we add a formula to the set already constructed, $\Delta_n \subseteq \Delta_{n+1}$, so since \subseteq is transitive, $\Delta_n \subseteq \Delta_m$ whenever $n \leq m$.) At each stage of the construction, we either add φ_n or $\neg\varphi_n$, and every formula appears (at least once) in the list of all φ_n . So, for every φ either $\varphi \in \Delta$ or $\neg\varphi \in \Delta$, so Δ is complete by definition.

Finally, we have to show, that Δ is Σ -consistent. To do this, we show that (a) if Δ were Σ -inconsistent, then some Δ_n would be Σ -inconsistent, and (b) all Δ_n are Σ -consistent.

So suppose Δ were Σ -inconsistent. Then $\Delta \vdash_{\Sigma} \perp$, i.e., there are $\varphi_1, \dots, \varphi_k \in \Delta$ such that $\Sigma \vdash \varphi_1 \rightarrow (\varphi_2 \rightarrow \dots (\varphi_k \rightarrow \perp) \dots)$. Since $\Delta = \bigcap_{n=0}^{\infty} \Delta_n$, each $\varphi_i \in \Delta_{n_i}$ for some n_i . Let n be the largest of these. Since $n_i \leq n$, $\Delta_{n_i} \subseteq \Delta_n$. So, all φ_i are in some Δ_n . This would mean $\Delta_n \vdash_{\Sigma} \perp$, i.e., Δ_n is Σ -inconsistent.

To show that each Δ_n is Σ -consistent, we use a simple induction on n . $\Delta_0 = \Gamma$, and we assumed Γ was Σ -consistent. So the claim holds for $n = 0$. Now suppose it holds for n , i.e., Δ_n is Σ -consistent. Δ_{n+1} is either $\Delta_n \cup \{\varphi_n\}$ or $\Delta_n \cup \{\neg\varphi_n\}$. In the first case, Δ_{n+1} is clearly Σ -consistent. However, by [Proposition 3.39\(3\)](#), either $\Delta_n \cup \{\varphi_n\}$ or $\Delta_n \cup \{\neg\varphi_n\}$ is consistent, so Δ_{n+1} is consistent in the other case as well. \square

[mod:com:lin:](#)
[cor:provability-characterization](#)

Corollary 4.4. *$\Gamma \vdash_{\Sigma} \varphi$ if and only if $\varphi \in \Delta$ for each complete Σ -consistent set Δ extending Γ (including when $\Gamma = \emptyset$, in which case we get another characterization of the modal system Σ .)*

Proof. Suppose $\Gamma \vdash_{\Sigma} \varphi$, and let Δ be any complete Σ -consistent set extending Γ . If $\varphi \notin \Delta$ then by maximality $\neg\varphi \in \Delta$ and so $\Delta \vdash_{\Sigma} \varphi$ (by monotony) and

$\Delta \vdash_{\Sigma} \neg\varphi$ (by reflexivity), and so Δ is inconsistent. Conversely if $\Gamma \not\vdash_{\Sigma} \varphi$, then $\Gamma \cup \{\neg\varphi\}$ is Σ -consistent, and by Lindenbaum's Lemma there is a complete consistent set Δ extending $\Gamma \cup \{\neg\varphi\}$. By consistency, $\varphi \notin \Delta$. \square

4.4 Modalities and Complete Consistent Sets

explanation

When we construct a model \mathfrak{M}^{Σ} whose set of worlds is given by the complete Σ -consistent sets Δ in some normal modal logic Σ , we will also need to define an accessibility relation R^{Σ} between such “worlds.” We want it to be the case that the accessibility relation (and the assignment V^{Σ}) are defined in such a way that $\mathfrak{M}^{\Sigma}, \Delta \Vdash \varphi$ iff $\varphi \in \Delta$. How should we do this?

mod:com:mod:
sec

Once the accessibility relation is defined, the definition of truth at a world ensures that $\mathfrak{M}^{\Sigma}, \Delta \Vdash \Box\varphi$ iff $\mathfrak{M}^{\Sigma}, \Delta' \Vdash \varphi$ for all Δ' such that $R^{\Sigma}\Delta\Delta'$. The proof that $\mathfrak{M}^{\Sigma}, \Delta \Vdash \varphi$ iff $\varphi \in \Delta$ requires that this is true in particular for **formulas** starting with a modal operator, i.e., $\mathfrak{M}^{\Sigma}, \Delta \Vdash \Box\varphi$ iff $\Box\varphi \in \Delta$. Combining this requirement with the definition of truth at a world for $\Box\varphi$ yields:

$$\Box\varphi \in \Delta \text{ iff } \varphi \in \Delta' \text{ for all } \Delta' \text{ with } R^{\Sigma}\Delta\Delta'$$

Consider the left-to-right direction: it says that if $\Box\varphi \in \Delta$, then $\varphi \in \Delta'$ for any φ and any Δ' with $R^{\Sigma}\Delta\Delta'$. If we stipulate that $R^{\Sigma}\Delta\Delta'$ iff $\varphi \in \Delta'$ for all $\Box\varphi \in \Delta$, then this holds. We can write the condition on the right of the “iff” more compactly as: $\{\varphi : \Box\varphi \in \Delta\} \subseteq \Delta'$.

So the question is: does this definition of R^{Σ} in fact guarantee that $\Box\varphi \in \Delta$ iff $\mathfrak{M}^{\Sigma}, \Delta \Vdash \Box\varphi$? Does it also guarantee that $\Diamond\varphi \in \Delta$ iff $\mathfrak{M}^{\Sigma}, \Delta \Vdash \Diamond\varphi$? The next few results will establish this.

Definition 4.5. If Γ is a set of **formulas**, let

$$\Box\Gamma = \{\Box\psi : \psi \in \Gamma\}$$

$$\Diamond\Gamma = \{\Diamond\psi : \psi \in \Gamma\}$$

and

$$\Box^{-1}\Gamma = \{\psi : \Box\psi \in \Gamma\}$$

$$\Diamond^{-1}\Gamma = \{\psi : \Diamond\psi \in \Gamma\}$$

In other words, $\Box\Gamma$ is Γ with \Box in front of every **formula** in Γ ; $\Box^{-1}\Gamma$ is all the \Box 'ed **formulas** of Γ with the initial \Box 's removed. This definition is not terribly important on its own, but will simplify the notation considerably.

Note that $\Box\Box^{-1}\Gamma \subseteq \Gamma$:

$$\Box\Box^{-1}\Gamma = \{\Box\psi : \Box\psi \in \Gamma\}$$

i.e., it's just the set of all those **formulas** of Γ that start with \Box .

mod:com:mod: **Lemma 4.6.** *lem:box1* If $\Gamma \vdash_{\Sigma} \varphi$ then $\Box\Gamma \vdash_{\Sigma} \Box\varphi$.

Proof. If $\Gamma \vdash_{\Sigma} \varphi$ then there are $\psi_1, \dots, \psi_k \in \Gamma$ such that $\Sigma \vdash \psi_1 \rightarrow (\psi_2 \rightarrow \dots (\psi_n \rightarrow \varphi) \dots)$. Since Σ is normal, by rule RK, $\Sigma \vdash \Box\psi_1 \rightarrow (\Box\psi_2 \rightarrow \dots (\Box\psi_n \rightarrow \Box\varphi) \dots)$, where obviously $\Box\psi_1, \dots, \Box\psi_k \in \Box\Gamma$. Hence, by definition, $\Box\Gamma \vdash_{\Sigma} \Box\varphi$. \square

mod:com:mod: **Lemma 4.7.** *lem:box2* If $\Box^{-1}\Gamma \vdash_{\Sigma} \varphi$ then $\Gamma \vdash_{\Sigma} \Box\varphi$.

Proof. Suppose $\Box^{-1}\Gamma \vdash_{\Sigma} \varphi$; then by Lemma 4.6, $\Box\Box^{-1}\Gamma \vdash \Box\varphi$. But since $\Box\Box^{-1}\Gamma \subseteq \Gamma$, also $\Gamma \vdash_{\Sigma} \Box\varphi$ by Monotony. \square

mod:com:mod: **Proposition 4.8.** *prop:box* If Γ is complete Σ -consistent, then $\Box\varphi \in \Gamma$ if and only if for every complete Σ -consistent Δ such that $\Box^{-1}\Gamma \subseteq \Delta$, it holds that $\varphi \in \Delta$.

Proof. Suppose Γ is complete Σ -consistent. The “only if” direction is easy: Suppose $\Box\varphi \in \Gamma$ and that $\Box^{-1}\Gamma \subseteq \Delta$. Since $\Box\varphi \in \Gamma$, $\varphi \in \Box^{-1}\Gamma \subseteq \Delta$, so $\varphi \in \Delta$.

For the “if” direction, we prove the contrapositive: Suppose $\Box\varphi \notin \Gamma$. Since Γ is complete Σ -consistent, it is deductively closed, and hence $\Gamma \not\vdash_{\Sigma} \Box\varphi$. By Lemma 4.7, $\Box^{-1}\Gamma \not\vdash_{\Sigma} \varphi$. By Proposition 3.39(2), $\Box^{-1}\Gamma \cup \{\neg\varphi\}$ is Σ -consistent. By Lindenbaum’s Lemma, there is a complete Σ -consistent set Δ such that $\Box^{-1}\Gamma \cup \{\neg\varphi\} \subseteq \Delta$. By consistency, $\varphi \notin \Delta$. \square

mod:com:mod: **Lemma 4.9.** *lem:box-iff-diamond* Suppose Γ and Δ are complete Σ -consistent. Then: $\Box^{-1}\Gamma \subseteq \Delta$ if and only if $\Diamond\Delta \subseteq \Gamma$.

Proof. “Only if” direction: Assume $\Box^{-1}\Gamma \subseteq \Delta$ and suppose $\Diamond\varphi \in \Diamond\Delta$ (i.e., $\varphi \in \Delta$). In order to show $\Diamond\varphi \in \Gamma$ it suffices to show $\Box\neg\varphi \notin \Gamma$ for then by maximality $\neg\Box\neg\varphi \in \Gamma$. Now, if $\Box\neg\varphi \in \Gamma$ then by hypothesis $\neg\varphi \in \Delta$, against the consistency of Δ (since $\varphi \in \Delta$). Hence $\Box\neg\varphi \notin \Gamma$, as required.

“If” direction: Assume $\Diamond\Delta \subseteq \Gamma$. We argue contrapositively: suppose $\varphi \notin \Delta$ in order to show $\Box\varphi \notin \Gamma$. If $\varphi \notin \Delta$ then by maximality $\neg\varphi \in \Delta$ and so by hypothesis $\Diamond\neg\varphi \in \Gamma$. But in a normal modal logic $\Diamond\neg\varphi$ is equivalent to $\neg\Box\varphi$, and if the latter is in Γ , by consistency $\Box\varphi \notin \Gamma$, as required. \square

mod:com:mod: **Proposition 4.10.** *prop:diamond* If Γ is complete Σ -consistent, then $\Diamond\varphi \in \Gamma$ if and only if for some complete Σ -consistent Δ such that $\Diamond\Delta \subseteq \Gamma$, it holds that $\varphi \in \Delta$.

Proof. Suppose Γ is complete Σ -consistent. $\Diamond\varphi \in \Gamma$ iff $\neg\Box\neg\varphi \in \Gamma$ by DUAL and closure. $\neg\Box\neg\varphi \in \Gamma$ iff $\Box\neg\varphi \notin \Gamma$ by Proposition 4.2(5) since Γ is complete Σ -consistent. By Proposition 4.8, $\Box\neg\varphi \notin \Gamma$ iff, for some complete Σ -consistent Δ with $\Box^{-1}\Gamma \subseteq \Delta$, $\neg\varphi \notin \Delta$. Now consider any such Δ . By Lemma 4.9, $\Box^{-1}\Gamma \subseteq \Delta$ iff $\Diamond\Delta \subseteq \Gamma$. Also, $\neg\varphi \notin \Delta$ iff $\varphi \in \Delta$ by Proposition 4.2(5). So $\Diamond\varphi \in \Gamma$ iff, for some complete Σ -consistent Δ with $\Diamond\Delta \subseteq \Gamma$, $\varphi \in \Delta$. \square

Problem 4.2. Show that if Γ is complete Σ -consistent, then $\Diamond\varphi \in \Gamma$ if and only if there is a complete Σ -consistent Δ such that $\Box^{-1}\Gamma \subseteq \Delta$ and $\varphi \in \Delta$. Do this without using Lemma 4.9.

4.5 Canonical Models

The *canonical model* for a modal system Σ is a specific model \mathfrak{M}^Σ in which the worlds are all complete Σ -consistent sets. Its accessibility relation R^Σ and valuation V^Σ are defined so as to guarantee that the formulas true at a world Δ are exactly the formulas making up Δ . mod:com:cmd:
sec

Definition 4.11. Let Σ be a normal modal logic. The *canonical model* for Σ is $\mathfrak{M}^\Sigma = \langle W^\Sigma, R^\Sigma, V^\Sigma \rangle$, where:

1. $\mathfrak{M}^\Sigma = \{\Delta : \Delta \text{ is complete } \Sigma\text{-consistent}\}$.
2. $R^\Sigma \Delta \Delta'$ holds if and only if $\Box^{-1} \Delta \subseteq \Delta'$.
3. $V^\Sigma(p) = \{\Delta : p \in \Delta\}$.

4.6 The Truth Lemma

The canonical model \mathfrak{M}^Σ is defined in such a way that $\mathfrak{M}^\Sigma, \Delta \Vdash \varphi$ iff $\varphi \in \Delta$. For propositional variables, the definition of V^Σ yields this directly. We have to verify that the equivalence holds for all formulas, however. We do this by induction. The inductive step involves proving the equivalence for formulas involving propositional operators (where we have to use Proposition 4.2) and the modal operators (where we invoke the results of section 4.4). mod:com:tru:
sec

Proposition 4.12 (Truth Lemma). *For every formula φ , $\mathfrak{M}^\Sigma, \Delta \Vdash \varphi$ if and only if $\varphi \in \Delta$.* mod:com:tru:
prop:truthlemma

Proof. By induction on φ .

1. $\varphi \equiv \perp$: $\mathfrak{M}^\Sigma, \Delta \not\Vdash \perp$ by Definition 1.6, and $\perp \notin \Delta$ by Proposition 4.2(3).
2. $\varphi \equiv \top$: $\mathfrak{M}^\Sigma, \Delta \Vdash \top$ by Definition 1.6, and $\top \in \Delta$ by Proposition 4.2(4).
3. $\varphi \equiv p$: $\mathfrak{M}^\Sigma, \Delta \Vdash p$ iff $\Delta \in V^\Sigma(p)$ by Definition 1.6. Also, $\Delta \in V^\Sigma(p)$ iff $p \in \Delta$ by definition of V^Σ .
4. $\varphi \equiv \neg\psi$: $\mathfrak{M}^\Sigma, \Delta \Vdash \neg\psi$ iff $\mathfrak{M}^\Sigma, \Delta \not\Vdash \psi$ (Definition 1.6) iff $\psi \notin \Delta$ (by inductive hypothesis) iff $\neg\psi \in \Delta$ (by Proposition 4.2(5)).
5. $\varphi \equiv \psi \wedge \chi$: $\mathfrak{M}^\Sigma, \Delta \Vdash \psi \wedge \chi$ iff $\mathfrak{M}^\Sigma, \Delta \Vdash \psi$ and $\mathfrak{M}^\Sigma, \Delta \Vdash \chi$ (by Definition 1.6) iff $\psi \in \Delta$ and $\chi \in \Delta$ (by inductive hypothesis) iff $\psi \wedge \chi \in \Delta$ (by Proposition 4.2(6)).
6. $\varphi \equiv \psi \vee \chi$: $\mathfrak{M}^\Sigma, \Delta \Vdash \psi \vee \chi$ iff $\mathfrak{M}^\Sigma, \Delta \Vdash \psi$ or $\mathfrak{M}^\Sigma, \Delta \Vdash \chi$ (by Definition 1.6) iff $\psi \in \Delta$ or $\chi \in \Delta$ (by inductive hypothesis) iff $\psi \vee \chi \in \Delta$ (by Proposition 4.2(7)).
7. $\varphi \equiv \psi \rightarrow \chi$: $\mathfrak{M}^\Sigma, \Delta \Vdash \psi \rightarrow \chi$ iff $\mathfrak{M}^\Sigma, \Delta \not\Vdash \psi$ or $\mathfrak{M}^\Sigma, \Delta \Vdash \chi$ (by Definition 1.6) iff $\psi \notin \Delta$ or $\chi \in \Delta$ (by inductive hypothesis) iff $\psi \rightarrow \chi \in \Delta$ (by Proposition 4.2(8)).

8. $\varphi \equiv \psi \leftrightarrow \chi$: $\mathfrak{M}^\Sigma, \Delta \Vdash \psi \leftrightarrow \chi$ iff either $\mathfrak{M}^\Sigma, \Delta \Vdash \psi$ and $\mathfrak{M}^\Sigma, \Delta \Vdash \chi$ or $\mathfrak{M}^\Sigma, \Delta \nVdash \psi$ and $\mathfrak{M}^\Sigma, \Delta \nVdash \chi$ (by [Definition 1.6](#)) iff either $\psi \in \Delta$ and $\chi \in \Delta$ or $\psi \notin \Delta$ and $\chi \notin \Delta$ (by inductive hypothesis) iff $\psi \leftrightarrow \chi \in \Delta$ (by [Proposition 4.2\(9\)](#)).
9. $\varphi \equiv \Box\psi$: First suppose that $\mathfrak{M}^\Sigma, \Delta \Vdash \Box\psi$. By [Definition 1.6](#), for every Δ' such that $R^\Sigma \Delta \Delta'$, $\mathfrak{M}^\Sigma, \Delta' \Vdash \psi$. By inductive hypothesis, for every Δ' such that $R^\Sigma \Delta \Delta'$, $\psi \in \Delta'$. By definition of R^Σ , for every Δ' such that $\Box^{-1}\Delta \subseteq \Delta'$, $\psi \in \Delta'$. By [Proposition 4.8](#), $\Box\psi \in \Delta$.
Now assume $\Box\psi \in \Delta$. Let $\Delta' \in W^\Sigma$ be such that $R^\Sigma \Delta \Delta'$, i.e., $\Box^{-1}\Delta \subseteq \Delta'$. Since $\Box\psi \in \Delta$, $\psi \in \Box^{-1}\Delta$. Consequently, $\psi \in \Delta'$. By inductive hypothesis, $\mathfrak{M}^\Sigma, \Delta' \Vdash \psi$. Since Δ' is arbitrary with $R^\Sigma \Delta \Delta'$, for all $\Delta' \in W^\Sigma$ such that $R^\Sigma \Delta \Delta'$, $\mathfrak{M}^\Sigma, \Delta' \Vdash \psi$. By [Definition 1.6](#), $\mathfrak{M}^\Sigma, \Delta \Vdash \Box\psi$.
10. $\varphi \equiv \Diamond\psi$: First suppose that $\mathfrak{M}^\Sigma, \Delta \Vdash \Diamond\psi$. By [Definition 1.6](#), for some Δ' such that $R^\Sigma \Delta \Delta'$, $\mathfrak{M}^\Sigma, \Delta' \Vdash \psi$. By inductive hypothesis, for some Δ' such that $R^\Sigma \Delta \Delta'$, $\psi \in \Delta'$. By definition of R^Σ , for some Δ' such that $\Box^{-1}\Delta \subseteq \Delta'$, $\psi \in \Delta'$. By [Proposition 4.10](#), for some Δ' such that $\Diamond\Delta' \subseteq \Delta$, $\psi \in \Delta'$. Since $\psi \in \Delta'$, $\Diamond\psi \in \Diamond\Delta'$, so $\Diamond\psi \in \Delta$.
Now assume $\Diamond\psi \in \Delta$. By [Proposition 4.10](#), there is a complete Σ -consistent $\Delta' \in W^\Sigma$ be such that $\Diamond\Delta' \subseteq \Delta$ and $\psi \in \Delta'$. By [Lemma 4.9](#), there is a $\Delta' \in W^\Sigma$ such that $\Box^{-1}\Delta \subseteq \Delta'$, and $\psi \in \Delta'$. By definition of R^Σ , $R^\Sigma \Delta \Delta'$, so there is a $\Delta' \in W^\Sigma$ such that $R^\Sigma \Delta \Delta'$ and $\psi \in \Delta'$. By [Definition 1.6](#), $\mathfrak{M}^\Sigma, \Delta \Vdash \Diamond\psi$.

□

Problem 4.3. Complete the proof of [Proposition 4.12](#).

4.7 Determination and Completeness for K

[mod:com:cmk:](#)
[sec](#)

We are now prepared to use the canonical model to establish completeness. Completeness follows from the fact that the [formulas](#) true in the canonical for Σ are exactly the Σ -[derivable](#) ones. Models with this property are said to *determine* Σ .

Definition 4.13. A model \mathfrak{M} *determines* a normal modal logic Σ precisely when $\mathfrak{M} \Vdash \varphi$ if and only if $\Sigma \vdash \varphi$, for all [formulas](#) φ .

[mod:com:cmk:](#)
[thm:determination](#)

Theorem 4.14 (Determination). $\mathfrak{M}^\Sigma \Vdash \varphi$ if and only if $\Sigma \vdash \varphi$.

Proof. If $\mathfrak{M}^\Sigma \Vdash \varphi$, then for every complete Σ -consistent Δ , we have $\mathfrak{M}^\Sigma, \Delta \Vdash \varphi$. Hence, by the Truth Lemma, $\varphi \in \Delta$ for every complete Σ -consistent Δ , whence by [Corollary 4.4](#) (with $\Gamma = \emptyset$), $\Sigma \vdash \varphi$.

Conversely, if $\Sigma \vdash \varphi$ then by [Proposition 4.2\(1\)](#), every complete Σ -consistent Δ contains φ , and hence by the Truth Lemma, $\mathfrak{M}^\Sigma, \Delta \Vdash \varphi$ for every $\Delta \in W^\Sigma$, i.e., $\mathfrak{M}^\Sigma \Vdash \varphi$. □

Since the canonical model for \mathbf{K} determines \mathbf{K} , we immediately have completeness of \mathbf{K} as a corollary:

Corollary 4.15. *The basic modal logic \mathbf{K} is complete with respect to the class of all models, i.e., if $\models \varphi$ then $\mathbf{K} \vdash \varphi$.* mod:com:cmk:
cor:Kcomplete

Proof. Contrapositively, if $\mathbf{K} \not\vdash \varphi$ then by Determination $\mathfrak{M}^{\mathbf{K}} \not\models \varphi$ and hence φ is not valid. \square

For the general case of completeness of a system Σ with respect to a class of models, e.g., of $\mathbf{KTB4}$ with respect to the class of reflexive, symmetric, transitive models, determination alone is not enough. We must also show that the canonical model for the system Σ is a member of the class, which does not follow obviously from the canonical model construction—nor is it always true!

4.8 Frame Completeness

The completeness theorem for \mathbf{K} can be extended to other modal systems, once we show that the canonical model for a given logic has the corresponding frame property. mod:com:fra:
sec

Theorem 4.16. *If a normal modal logic Σ contains one of the formulas on the left-hand side of table 4.1, then the canonical model for Σ has the corresponding property on the right-hand side.* mod:com:fra:
thm:completeframeprops

<i>If Σ contains ...</i>	<i>... the canonical model for Σ is:</i>
D: $\Box\varphi \rightarrow \Diamond\varphi$	serial;
T: $\Box\varphi \rightarrow \varphi$	reflexive;
B: $\varphi \rightarrow \Box\Diamond\varphi$	symmetric;
4: $\Box\varphi \rightarrow \Box\Box\varphi$	transitive;
5: $\Diamond\varphi \rightarrow \Box\Diamond\varphi$	euclidean.

Table 4.1: Basic correspondence facts.

Proof. We take each of these up in turn. mod:com:fra:
tab:correspondencetable

Suppose Σ contains D, and let $\Delta \in W^\Sigma$; we need to show that there is a Δ' such that $R^\Sigma \Delta \Delta'$. It suffices to show that $\Box^{-1}\Delta$ is Σ -consistent, for then by Lindenbaum's Lemma, there is a complete Σ -consistent set $\Delta' \supseteq \Box^{-1}\Delta$, and by definition of R^Σ we have $R^\Sigma \Delta \Delta'$. So, suppose for contradiction that $\Box^{-1}\Delta$ is *not* Σ -consistent, i.e., $\Box^{-1}\Delta \vdash_\Sigma \perp$. By Lemma 4.7, $\Delta \vdash_\Sigma \Box\perp$, and since Σ contains D, also $\Delta \vdash_\Sigma \Diamond\perp$. But Σ is normal, so $\Sigma \vdash \neg\Diamond\perp$ (Proposition 3.7), whence also $\Delta \vdash_\Sigma \neg\Diamond\perp$, against the consistency of Δ .

Now suppose Σ contains T, and let $\Delta \in W^\Sigma$. We want to show $R^\Sigma \Delta \Delta$, i.e., $\Box^{-1}\Delta \subseteq \Delta$. But if $\Box\varphi \in \Delta$ then by T also $\varphi \in \Delta$, as desired.

Now suppose Σ contains B, and suppose $R^\Sigma \Delta \Delta'$ for $\Delta, \Delta' \in W^\Sigma$. We need to show that $R^\Sigma \Delta' \Delta$, i.e., $\Box^{-1}\Delta' \subseteq \Delta$. By Lemma 4.9, this is equivalent

to $\diamond\Delta \subseteq \Delta'$. So suppose $\varphi \in \Delta$. By B, also $\Box\diamond\varphi \in \Delta$. By the hypothesis that $R^\Sigma\Delta\Delta'$, we have that $\Box^{-1}\Delta \subseteq \Delta'$, and hence $\diamond\varphi \in \Delta'$, as required.

Now suppose Σ contains 4, and suppose $R^\Sigma\Delta_1\Delta_2$ and $R^\Sigma\Delta_2\Delta_3$. We need to show $R^\Sigma\Delta_1\Delta_3$. From the hypothesis we have both $\Box^{-1}\Delta_1 \subseteq \Delta_2$ and $\Box^{-1}\Delta_2 \subseteq \Delta_3$. In order to show $R^\Sigma\Delta_1\Delta_3$ it suffices to show $\Box^{-1}\Delta_1 \subseteq \Delta_3$. So let $\psi \in \Box^{-1}\Delta_1$, i.e., $\Box\psi \in \Delta_1$. By 4, also $\Box\Box\psi \in \Delta_1$ and by hypothesis we get, first, that $\Box\psi \in \Delta_2$ and, second, that $\psi \in \Delta_3$, as desired.

Now suppose Σ contains 5, suppose $R^\Sigma\Delta_1\Delta_2$ and $R^\Sigma\Delta_1\Delta_3$. We need to show $R^\Sigma\Delta_2\Delta_3$. The first hypothesis gives $\Box^{-1}\Delta_1 \subseteq \Delta_2$, and the second hypothesis is equivalent to $\diamond\Delta_3 \subseteq \Delta_2$, by [Lemma 4.9](#). To show $R^\Sigma\Delta_2\Delta_3$, by [Lemma 4.9](#), it suffices to show $\diamond\Delta_3 \subseteq \Delta_2$. So let $\diamond\varphi \in \diamond\Delta_3$, i.e., $\varphi \in \Delta_3$. By the second hypothesis $\diamond\varphi \in \Delta_1$ and by 5, $\Box\diamond\varphi \in \Delta_1$ as well. But now the first hypothesis gives $\diamond\varphi \in \Delta_2$, as desired. \square

As a corollary we obtain completeness results for a number of systems. For instance, we know that **S5** = **KT5** = **KTB4** is complete with respect to the class of all reflexive euclidean models, which is the same as the class of all reflexive, symmetric and transitive models.

*mod:com:fra:
thm:generaldet*

Theorem 4.17. *Let $\mathcal{C}_D, \mathcal{C}_T, \mathcal{C}_B, \mathcal{C}_4$, and \mathcal{C}_5 be the class of all serial, reflexive, symmetric, transitive, and euclidean models (respectively). Then for any schemas $\varphi_1, \dots, \varphi_n$ among D, T, B, 4, and 5, the system $\mathbf{K}\varphi_1 \dots \varphi_n$ is determined by the class of models $\mathcal{C} = \mathcal{C}_{\varphi_1} \cap \dots \cap \mathcal{C}_{\varphi_n}$.*

Proposition 4.18. *Let Σ be a normal modal logic; then:*

*mod:com:fra:
prop:anotherfive-a*

1. *If Σ contains the schema $\diamond\varphi \rightarrow \Box\varphi$ then the canonical model for Σ is partially functional.*
2. *If Σ contains the schema $\diamond\varphi \leftrightarrow \Box\varphi$ then the canonical model for Σ is functional.*
3. *If Σ contains the schema $\Box\Box\varphi \rightarrow \Box\varphi$ then the canonical model for Σ is weakly dense.*

(see [table 2.2](#) for definitions of these frame properties).

Proof. 1. suppose that Σ contains the schema $\diamond\varphi \rightarrow \Box\varphi$, to show that R^Σ is partially functional we need to prove that for any $\Delta_1, \Delta_2, \Delta_3 \in W^\Sigma$, if $R^\Sigma\Delta_1\Delta_2$ and $R^\Sigma\Delta_1\Delta_3$ then $\Delta_2 = \Delta_3$. Since $R^\Sigma\Delta_1\Delta_2$ we have $\Box^{-1}\Delta_1 \subseteq \Delta_2$ and since $R^\Sigma\Delta_1\Delta_3$ also $\Box^{-1}\Delta_1 \subseteq \Delta_3$. The identity $\Delta_2 = \Delta_3$ will follow if we can establish the two inclusions $\Delta_2 \subseteq \Delta_3$ and $\Delta_3 \subseteq \Delta_2$. For the first inclusion, let $\varphi \in \Delta_2$; then $\diamond\varphi \in \Delta_1$, and by the schema and deductive closure of Δ_1 also $\Box\varphi \in \Delta_1$, whence by the hypothesis that $R^\Sigma\Delta_1\Delta_3$, $\varphi \in \Delta_3$. The second inclusion is similar.

2. This follows immediately from part (1) and the seriality proof in [Theorem 4.16](#).

3. Suppose Σ contains the schema $\Box\Box\varphi \rightarrow \Box\varphi$ and to show that R^Σ is weakly dense, let $R^\Sigma\Delta_1\Delta_2$. We need to show that there is a complete Σ -consistent set Δ_3 such that $R^\Sigma\Delta_1\Delta_3$ and $R^\Sigma\Delta_3\Delta_2$. Let:

$$\Gamma = \Box^{-1}\Delta_1 \cup \Diamond\Delta_2.$$

It suffices to show that Γ is Σ -consistent, for then by Lindenbaum's Lemma it can be extended to a complete Σ -consistent set Δ_3 such that $\Box^{-1}\Delta_1 \subseteq \Delta_3$ and $\Diamond\Delta_2 \subseteq \Delta_3$, i.e., $R^\Sigma\Delta_1\Delta_3$ and $R^\Sigma\Delta_3\Delta_2$ (by Lemma 4.9).

Suppose for contradiction that Γ is not consistent. Then there are formulas $\Box\varphi_1, \dots, \Box\varphi_n \in \Delta_1$ and $\psi_1, \dots, \psi_m \in \Delta_2$ such that

$$\varphi_1, \dots, \varphi_n, \Diamond\psi_1, \dots, \Diamond\psi_m \vdash_\Sigma \perp.$$

Since $\Diamond(\psi_1 \wedge \dots \wedge \psi_m) \rightarrow (\Diamond\psi_1 \wedge \dots \wedge \Diamond\psi_m)$ is derivable in every normal modal logic, we argue as follows, contradicting the consistency of Δ_2 :

$$\begin{aligned} & \varphi_1, \dots, \varphi_n, \Diamond\psi_1, \dots, \Diamond\psi_m \vdash_\Sigma \perp \\ & \Rightarrow \varphi_1, \dots, \varphi_n \vdash_\Sigma (\Diamond\psi_1 \wedge \dots \wedge \Diamond\psi_m) \rightarrow \perp, & \text{deduction theorem;} \\ & \Rightarrow \varphi_1, \dots, \varphi_n \vdash_\Sigma \Diamond(\psi_1 \wedge \dots \wedge \psi_m) \rightarrow \perp, & \Sigma \text{ is normal;} \\ & \Rightarrow \varphi_1, \dots, \varphi_n \vdash_\Sigma \Box\neg(\psi_1 \wedge \dots \wedge \psi_m), & \text{PL;} \\ & \Rightarrow \Box\varphi_1, \dots, \Box\varphi_n \vdash_\Sigma \Box\Box\neg(\psi_1 \wedge \dots \wedge \psi_m), & \text{Lemma 4.6;} \\ & \Rightarrow \Box\varphi_1, \dots, \Box\varphi_n \vdash_\Sigma \Box\neg(\psi_1 \wedge \dots \wedge \psi_m), & \text{by the schema;} \\ & \Rightarrow \Delta_1 \vdash_\Sigma \Box\neg(\psi_1 \wedge \dots \wedge \psi_m), & \text{Monotony;} \\ & \Rightarrow \Box\neg(\psi_1 \wedge \dots \wedge \psi_m) \in \Delta_1, & \text{deductive closure;} \\ & \Rightarrow \neg(\psi_1 \wedge \dots \wedge \psi_m) \in \Delta_2, & \text{since } R^\Sigma\Delta_1\Delta_2. \end{aligned}$$

□

On the strength of these examples, one might think that every system Σ of modal logic is *complete*, in the sense that it proves every formula which is valid in every frame in which every theorem of Σ is valid. Unfortunately, there are many systems that are not complete in this sense.

Chapter 5

Filtrations and Decidability

5.1 Introduction

mod:fil:int:
sec

One important question about a logic is always whether it is decidable, i.e., if there is an effective procedure which will answer the question “is this formula valid.” Propositional logic is decidable: we can effectively test if a formula is a tautology by constructing a truth table, and for a given formula, the truth table is finite. But we can’t obviously test if a modal formula is true in all models, for there are infinitely many of them. We can list all the finite models relevant to a given formula, since only the assignment of subsets of worlds to propositional variables which actually occur in the formula are relevant. If the accessibility relation is fixed, the possible different assignments $V(p)$ are just all the subsets of W , and if $|W| = n$ there are 2^n of those. If our formula φ contains m propositional variables there are then 2^{nm} different models with n worlds. For each one, we can test if φ is true at all worlds, simply by computing the truth value of φ in each. Of course, we also have to check all possible accessibility relations, but there are only finitely many relations on n worlds as well (specifically, the number of subsets of $W \times W$, i.e., 2^{n^2}).

If we are not interested in the logic **K**, but a logic defined by some class of models (e.g., the reflexive transitive models), we also have to be able to test if the accessibility relation is of the right kind. We can do that whenever the frames we are interested in are definable by modal formulas (e.g., by testing if T and 4 valid in the frame). So, the idea would be to run through all the finite frames, test each one if it is a frame in the class we’re interested in, then list all the possible models on that frame and test if φ is true in each. If not, stop: φ is not valid in the class of models of interest.

There is a problem with this idea: we don’t know when, if ever, we can stop looking. If the formula has a finite countermodel, our procedure will find it. But if it has no finite countermodel, we won’t get an answer. The formula may be valid (no countermodels at all), or it have only an infinite countermodel, which we’ll never look at. This problem can be overcome if we can show that every formula that has a countermodel has a finite countermodel. If this is the

case we say the logic has the *finite model property*.

But how would we show that a logic has the finite model property? One way of doing this would be to find a way to turn an infinite (counter)model of φ into a finite one. If that can be done, then whenever there is a model in which φ is not true, then the resulting finite model also makes φ not true. That finite model will show up on our list of all finite models, and we will eventually determine, for every formula that is not valid, that it isn't. Our procedure won't terminate if the formula is valid. If we can show in addition that there is some maximum size that the finite model our procedure provides can have, and that this maximum size depends only on the formula φ , we will have a size up to which we have to test finite models in our search for countermodels. If we haven't found a countermodel by then, there are none. Then our procedure will, in fact, decide the question "is φ valid?" for any formula φ .

A strategy that often works for turning infinite structures into finite structures is that of "identifying" elements of the structure which behave the same way in relevant respects. If there are infinitely many worlds in \mathfrak{M} that behave the same in relevant respects, then we might hope that there are only finitely many "classes" of such worlds. In other words, we partition the set of worlds in the right way. Each partition contains infinitely many worlds, but there are only finitely many partitions. Then we define a new model \mathfrak{M}^* where the worlds are the partitions. Finitely many partitions in the old model give us finitely many worlds in the new model, i.e., a finite model. Let's call the partition a world w is in $[w]$. We'll want it to be the case that $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \varphi$, since we want the new model to be a countermodel to φ if the old one was. This requires that we define the partition, as well as the accessibility relation of \mathfrak{M}^* in the right way.

To see how this would go, first imagine we have no accessibility relation. $\mathfrak{M}, w \Vdash \Box\psi$ iff for some $v \in W$, $\mathfrak{M}, v \Vdash \Box\psi$, and the same for \mathfrak{M}^* , except with $[w]$ and $[v]$. As a first idea, let's say that two worlds u and v are equivalent (belong to the same partition) if they agree on all propositional variables in \mathfrak{M} , i.e., $\mathfrak{M}, u \Vdash p$ iff $\mathfrak{M}, v \Vdash p$. Let $V^*(p) = \{[w] : \mathfrak{M}, w \Vdash p\}$. Our aim is to show that $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \varphi$. Obviously, we'd prove this by induction: The base case would be $\varphi \equiv p$. First suppose $\mathfrak{M}, w \Vdash p$. Then $[w] \in V^*(p)$ by definition, so $\mathfrak{M}^*, [w] \Vdash p$. Now suppose that $\mathfrak{M}^*, [w] \Vdash p$. That means that $[w] \in V^*(p)$, i.e., for some v equivalent to w , $\mathfrak{M}, v \Vdash p$. But " w equivalent to v " means " w and v make all the same propositional variables true," so $\mathfrak{M}, w \Vdash p$. Now for the inductive step, e.g., $\varphi \equiv \neg\psi$. Then $\mathfrak{M}, w \Vdash \neg\psi$ iff $\mathfrak{M}, w \not\Vdash \psi$ iff $\mathfrak{M}^*, [w] \not\Vdash \psi$ (by inductive hypothesis) iff $\mathfrak{M}^*, [w] \Vdash \neg\psi$. Similarly for the other non-modal operators. It also works for \Box : suppose $\mathfrak{M}^*, [w] \Vdash \Box\psi$. That means that for every $[u]$, $\mathfrak{M}^*, [u] \Vdash \psi$. By inductive hypothesis, for every u , $\mathfrak{M}, u \Vdash \psi$. Consequently, $\mathfrak{M}, w \Vdash \Box\psi$.

In the general case, where we have to also define the accessibility relation for \mathfrak{M}^* , things are more complicated. We'll call a model \mathfrak{M}^* a *filtration* if its accessibility relation R^* satisfies the conditions required to make the inductive proof above go through. Then any filtration \mathfrak{M}^* will make φ true at $[w]$ iff \mathfrak{M} makes φ true at w . However, now we also have to show that there are

filtrations, i.e., we can define R^* so that it satisfies the required conditions. In order for this to work, however, we have to require that worlds u, v count as equivalent not just when they agree on all **propositional variables**, but on all sub-formulas of φ . Since φ has only finitely many sub-formulas, this will still guarantee that the filtration is finite. There is not just one way to define a filtration, and in order to make sure that the accessibility relation of the filtration satisfies the required properties (e.g., reflexive, transitive, etc.) we have to be inventive with the definition of R^* .

5.2 Preliminaries

mod:fil:pre:sec Filtrations allow us to establish the decidability of our systems of modal logic by showing that they have the *finite model property*, i.e., that any **formula** that is true (false) in a model is also true (false) in a *finite* model. Filtrations are defined relative to sets of **formulas** which are closed under subformulas.

mod:fil:pre: defn:modallyclosed **Definition 5.1.** A set Γ of **formulas** is *closed under subformulas* if it contains every subformula of a **formula** in Γ . Further, Γ is *modally closed* if it is closed under subformulas and moreover $\varphi \in \Gamma$ implies $\Box\varphi, \Diamond\varphi \in \Gamma$.

For instance, given a **formula** φ , the set of all its sub-formulas is closed under sub-formulas. When we're defining a filtration of a model through the set of sub-formulas of φ , it will have the property we're after: it makes φ true (false) iff the original model does.

The set of worlds of a filtration of \mathfrak{M} through Γ is defined as the set of all equivalence classes of the following equivalence relation.

Definition 5.2. Let $\mathfrak{M} = \langle W, R, V \rangle$ and suppose Γ is closed under sub-formulas. Define a relation \equiv on W to hold of any two worlds that make the same **formulas** from Γ true, i.e.:

$$u \equiv v \quad \text{if and only if} \quad \forall \varphi \in \Gamma : \mathfrak{M}, u \Vdash \varphi \Leftrightarrow \mathfrak{M}, v \Vdash \varphi.$$

The equivalence class $[w]_{\equiv}$ of a world w , or $[w]$ for short, is the set of all worlds \equiv -equivalent to w :

$$[w] = \{v : v \equiv w\}.$$

Proposition 5.3. *Given \mathfrak{M} and Γ , \equiv as defined above is an equivalence relation, i.e., it is reflexive, symmetric, and transitive.*

Proof. The relation \equiv is reflexive, since w makes exactly the same **formulas** from Γ true as itself. It is symmetric since if u makes the same **formulas** from Γ true as v , the same holds for v and u . It is also transitive, since if u makes the same **formulas** from Γ true as v , and v as w , then u makes the same **formulas** from Γ true as w . \square

The relation \equiv , like any equivalence relation, divides W into *partitions*, i.e., subsets of W which are pairwise disjoint, and together cover all of W . Every $w \in W$ is an **element** of one of the partitions, namely of $[w]$, since $w \equiv w$. So the partitions $[w]$ cover all of W . They are pairwise disjoint, for if $u \in [w]$ and $u \in [v]$, then $u \equiv w$ and $u \equiv v$, and by symmetry and transitivity, $w \equiv v$, and so $[w] = [v]$.

5.3 Filtrations

Rather than define “the” filtration of \mathfrak{M} through Γ , we define when a model \mathfrak{M}^* counts as a filtration of \mathfrak{M} . All filtrations have the same set of worlds W^* and the same valuation V^* . But different filtrations may have different accessibility relations R^* . To count as a filtration, R^* has to satisfy a number of conditions, however. These conditions are exactly what we’ll require to prove the main result, namely that $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \varphi$, provided $\varphi \in \Gamma$. mod:fil:fil:
sec

Definition 5.4. Let Γ be closed under subformulas and $\mathfrak{M} = \langle W, R, V \rangle$. A *filtration of \mathfrak{M} through Γ* is any model $\mathfrak{M}^* = \langle W^*, R^*, V^* \rangle$, where: mod:fil:fil:
defn:filtration

1. $W^* = \{[w] : w \in W\}$;
2. For any $u, v \in W$: mod:fil:fil:
defn:filtration-R
 - a) If Ruv then $R^*[u][v]$;
 - b) If $R^*[u][v]$ then for any $\Box\varphi \in \Gamma$, if $\mathfrak{M}, u \Vdash \Box\varphi$ then $\mathfrak{M}, v \Vdash \varphi$;
 - c) If $R^*[u][v]$ then for any $\Diamond\varphi \in \Gamma$, if $\mathfrak{M}, v \Vdash \varphi$ then $\mathfrak{M}, u \Vdash \Diamond\varphi$. mod:fil:fil:
defn:filtration-R1
defn:filtration-R2
mod:fil:fil:
defn:filtration-R3
3. $V^*(p) = \{[u] : u \in V(p)\}$.

It’s worthwhile thinking about what $V^*(p)$ is: the set consisting of the equivalence classes $[w]$ of all worlds w where p is true in \mathfrak{M} . On the one hand, if $w \in V(p)$, then $[w] \in V^*(p)$ by that definition. However, it is not necessarily the case that if $[w] \in V^*(p)$, then $w \in V(p)$. If $[w] \in V^*(p)$ we are only guaranteed that $[w] = [u]$ for *some* $u \in V(p)$. Of course, $[w] = [u]$ means that $w \equiv u$. So, when $[w] \in V^*(p)$ we can (only) conclude that $w \equiv u$ for some $u \in V(p)$.

Theorem 5.5. *If \mathfrak{M}^* is a filtration of \mathfrak{M} through Γ , then for every $\varphi \in \Gamma$ and $w \in W$, we have $\mathfrak{M}, w \Vdash \varphi$ if and only if $\mathfrak{M}^*, [w] \Vdash \varphi$.* mod:fil:fil:
thm:filtrations

Proof. By induction on φ , using the fact that Γ is closed under subformulas. Since $\varphi \in \Gamma$ and Γ is closed under sub-formulas, all sub-formulas of φ are also $\in \Gamma$. Hence in each inductive step, the induction hypothesis applies to the sub-formulas of φ .

1. $\varphi \equiv \perp$: Neither $\mathfrak{M}, w \Vdash \varphi$ nor $\mathfrak{M}^*, [w] \Vdash \varphi$.

2. $\varphi \equiv \top$: Both $\mathfrak{M}, w \Vdash \varphi$ and $\mathfrak{M}^*, w \Vdash \varphi$.
3. $\varphi \equiv p$: The left-to-right direction is immediate, as $\mathfrak{M}, w \Vdash \varphi$ only if $w \in V(p)$, which implies $[w] \in V^*(p)$, i.e., $\mathfrak{M}^*, [w] \Vdash \varphi$. Conversely, suppose $\mathfrak{M}^*, [w] \Vdash \varphi$, i.e., $[w] \in V^*(p)$. Then for some $v \in V(p)$, $w \equiv v$. Of course then also $\mathfrak{M}, v \Vdash p$. Since $w \equiv v$, w and v make the same formulas from Γ true. Since by assumption $p \in \Gamma$ and $\mathfrak{M}, v \Vdash p$, $\mathfrak{M}, w \Vdash \varphi$.
4. $\varphi \equiv \neg\psi$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \not\Vdash \psi$. By induction hypothesis, $\mathfrak{M}, w \not\Vdash \psi$ iff $\mathfrak{M}^*, [w] \not\Vdash \psi$. Finally, $\mathfrak{M}^*, [w] \not\Vdash \psi$ iff $\mathfrak{M}^*, [w] \Vdash \varphi$.
5. $\varphi \equiv (\psi \wedge \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \Vdash \psi$ and $\mathfrak{M}, w \Vdash \chi$. By induction hypothesis, $\mathfrak{M}, w \Vdash \psi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$, and $\mathfrak{M}, w \Vdash \chi$ iff $\mathfrak{M}^*, [w] \Vdash \chi$. And $\mathfrak{M}^*, [w] \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$ and $\mathfrak{M}^*, [w] \Vdash \chi$.
6. $\varphi \equiv (\psi \vee \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \Vdash \psi$ or $\mathfrak{M}, w \Vdash \chi$. By induction hypothesis, $\mathfrak{M}, w \Vdash \psi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$, and $\mathfrak{M}, w \Vdash \chi$ iff $\mathfrak{M}^*, [w] \Vdash \chi$. And $\mathfrak{M}^*, [w] \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$ or $\mathfrak{M}^*, [w] \Vdash \chi$.
7. $\varphi \equiv (\psi \rightarrow \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \not\Vdash \psi$ or $\mathfrak{M}, w \Vdash \chi$. By induction hypothesis, $\mathfrak{M}, w \not\Vdash \psi$ iff $\mathfrak{M}^*, [w] \not\Vdash \psi$, and $\mathfrak{M}, w \Vdash \chi$ iff $\mathfrak{M}^*, [w] \Vdash \chi$. And $\mathfrak{M}^*, [w] \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \not\Vdash \psi$ or $\mathfrak{M}^*, [w] \Vdash \chi$.
8. $\varphi \equiv (\psi \leftrightarrow \chi)$: $\mathfrak{M}, w \Vdash \varphi$ iff $\mathfrak{M}, w \Vdash \psi$ and $\mathfrak{M}, w \Vdash \chi$, or $\mathfrak{M}, w \not\Vdash \psi$ and $\mathfrak{M}, w \not\Vdash \chi$. By induction hypothesis, $\mathfrak{M}, w \Vdash \psi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$, and $\mathfrak{M}, w \Vdash \chi$ iff $\mathfrak{M}^*, [w] \Vdash \chi$. And $\mathfrak{M}^*, [w] \Vdash \varphi$ iff $\mathfrak{M}^*, [w] \Vdash \psi$ and $\mathfrak{M}^*, [w] \Vdash \chi$, or $\mathfrak{M}^*, [w] \not\Vdash \psi$ and $\mathfrak{M}^*, [w] \not\Vdash \chi$.
9. $\varphi \equiv \Box\psi$: Suppose $\mathfrak{M}, w \Vdash \varphi$; to show that $\mathfrak{M}^*, [w] \Vdash \varphi$, let v be such that $R^*[w][v]$. From [Definition 5.4\(2b\)](#), we have that $\mathfrak{M}, v \Vdash \psi$, and by inductive hypothesis $\mathfrak{M}^*, [v] \Vdash \psi$. Since v was arbitrary, $\mathfrak{M}^*, [w] \Vdash \varphi$ follows.
Conversely, suppose $\mathfrak{M}^*, [w] \Vdash \varphi$ and let v be arbitrary such that Rwv . From [Definition 5.4\(2a\)](#), we have $R^*[w][v]$, so that $\mathfrak{M}^*, [v] \Vdash \psi$; by inductive hypothesis $\mathfrak{M}, v \Vdash \psi$, and since v was arbitrary, $\mathfrak{M}, w \Vdash \varphi$.
10. $\varphi \equiv \Diamond\psi$: Suppose $\mathfrak{M}, w \Vdash \varphi$. Then for some $v \in W$, Rwv and $\mathfrak{M}, v \Vdash \psi$. By inductive hypothesis $\mathfrak{M}^*, [v] \Vdash \psi$, and by [Definition 5.4\(2a\)](#), we have $R^*[w][v]$. Thus, $\mathfrak{M}^*, [w] \Vdash \varphi$.
Now suppose $\mathfrak{M}^*, [w] \Vdash \varphi$. Then for some $[v] \in W^*$ with $R^*[w][v]$, $\mathfrak{M}^*, [v] \Vdash \psi$. By inductive hypothesis $\mathfrak{M}, v \Vdash \psi$. By [Definition 5.4\(2c\)](#), we have that $\mathfrak{M}, w \Vdash \varphi$.

□

Problem 5.1. Complete the proof of [Theorem 5.5](#)

What holds for truth at worlds in a model also holds for truth in a model and validity in a class of models.

Corollary 5.6. *Let Γ be closed under subformulas. Then:*

1. *If \mathfrak{M}^* is a filtration of \mathfrak{M} through Γ then for any $\varphi \in \Gamma$: $\mathfrak{M} \Vdash \varphi$ if and only if $\mathfrak{M}^* \Vdash \varphi$.*
2. *If \mathcal{C} is a class of models and $\Gamma(\mathcal{C})$ is the class of Γ -filtrations of models in \mathcal{C} , then any formula $\varphi \in \Gamma$ is valid in \mathcal{C} if and only if it is valid in $\Gamma(\mathcal{C})$.*

5.4 Examples of Filtrations

We have not yet shown that there are any filtrations. But indeed, for any model \mathfrak{M} , there are many filtrations of \mathfrak{M} through Γ . We identify two, in particular: the finest and coarsest filtrations. Filtrations of the same models will differ in their accessibility relation (as Definition 5.4 stipulates directly what W^* and V^* should be). The finest filtration will have as few related worlds as possible, whereas the coarsest will have as many as possible. mod:fil:exf:
sec

Definition 5.7. Where Γ is closed under subformulas, the *finest* filtration \mathfrak{M}^* of a model \mathfrak{M} is defined by putting:

$$R^*[u][v] \text{ if and only if } \exists u' \in [u] \exists v' \in [v] : Ru'v'.$$

Proposition 5.8. *The finest filtration \mathfrak{M}^* is indeed a filtration.* mod:fil:exf:
prop:finest

Proof. We need to check that R^* , so defined, satisfies Definition 5.4(2). We check the three conditions in turn.

If Ruv then since $u \in [u]$ and $v \in [v]$, also $R^*[u][v]$, so (2a) is satisfied.

For (2b), suppose $\Box\varphi \in \Gamma$, $R^*[u][v]$, and $\mathfrak{M}, u \Vdash \Box\varphi$. By definition of R^* , there are $u' \equiv u$ and $v' \equiv v$ such that $Ru'v'$. Since u and u' agree on Γ , also $\mathfrak{M}, u' \Vdash \Box\varphi$, so that $\mathfrak{M}, v' \Vdash \varphi$. By closure of Γ under sub-formulas, v and v' agree on φ , so $\mathfrak{M}, v \Vdash \varphi$, as desired.

To verify (2c), suppose $\Diamond\varphi \in \Gamma$, $R^*[u][v]$, and $\mathfrak{M}, v \Vdash \varphi$. By definition of R^* , there are $u' \equiv u$ and $v' \equiv v$ such that $Ru'v'$. Since v and v' agree on Γ , and Γ is closed under sub-formulas, also $\mathfrak{M}, v' \Vdash \varphi$, so that $\mathfrak{M}, u' \Vdash \Diamond\varphi$. Since u and u' also agree on Γ , $\mathfrak{M}, u \Vdash \Diamond\varphi$. \square

Problem 5.2. Complete the proof of Proposition 5.8.

Definition 5.9. Where Γ is closed under subformulas, the *coarsest* filtration \mathfrak{M}^* of a model \mathfrak{M} is defined by putting $R^*[u][v]$ if and only if *both* of the following conditions are met:

1. If $\Box\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Box\varphi$ then $\mathfrak{M}, v \Vdash \varphi$;
2. If $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \varphi$ then $\mathfrak{M}, u \Vdash \Diamond\varphi$.

mod:fil:exf:
defn:coarsest-Box
mod:fil:exf:
defn:coarsest-Diamond

Proposition 5.10. *The coarsest filtration \mathfrak{M}^* is indeed a filtration.*

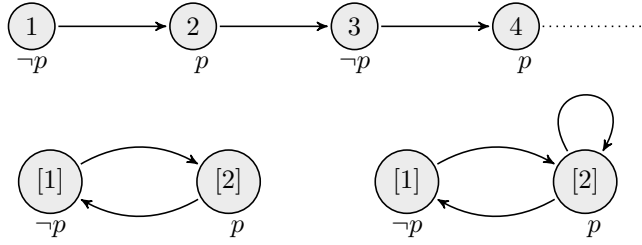


Figure 5.1: An infinite model and its filtrations.

mod:fil:exf:
fig:ex-filtration

Proof. Given the definition of R^* , the only condition that is left to verify is the implication from Ruv to $R^*[u][v]$. So assume Ruv . Suppose $\Box\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Box\varphi$; then obviously $\mathfrak{M}, v \Vdash \varphi$, and (1) is satisfied. Suppose $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \varphi$. Then $\mathfrak{M}, u \Vdash \Diamond\varphi$ since Ruv , and (2) is satisfied. \square

Example 5.11. Let $W = \mathbb{Z}^+$, Rnm iff $m = n + 1$, and $V(p) = \{2n : n \in \mathbb{N}\}$. The model $\mathfrak{M} = \langle W, R, V \rangle$ is depicted in Figure 5.1. The worlds are 1, 2, etc.; each world can access exactly one other world—its successor, and p is true at all and only the even numbers.

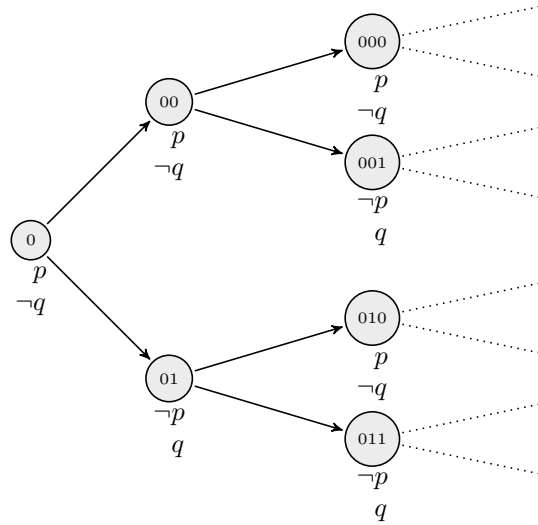
Now let Γ be the set of sub-formulas of $\Box p \rightarrow p$, i.e., $\{p, \Box p, \Box p \rightarrow p\}$. p is true at all and only the even numbers, $\Box p$ is true at all and only the odd numbers, so $\Box p \rightarrow p$ is true at all and only the even numbers. In other words, every odd number makes $\Box p$ true and p and $\Box p \rightarrow p$ false; every even number makes p and $\Box p \rightarrow p$ true, but $\Box p$ false. So $W^* = \{[1], [2]\}$, where $[1] = \{1, 3, 5, \dots\}$ and $[2] = \{2, 4, 6, \dots\}$. Since $2 \in V(p)$, $[2] \in V^*(p)$; since $1 \notin V(p)$, $[1] \notin V^*(p)$. So $V^*(p) = \{[2]\}$.

Any filtration based on W^* must have an accessibility relation that includes $\langle [1], [2] \rangle, \langle [2], [1] \rangle$: since $R12$, we must have $R^*[1][2]$ by Definition 5.4(2a), and since $R23$ we must have $R^*[2][3]$, and $[3] = [1]$. It cannot include $\langle [1], [1] \rangle$: if it did, we'd have $R^*[1][1]$, $\mathfrak{M}, 1 \Vdash \Box p$ but $\mathfrak{M}, 1 \not\Vdash p$, contradicting (2a). Nothing requires or rules out that $R^*[2][2]$. So, there are two possible filtrations of \mathfrak{M} , corresponding to the two accessibility relations

$$\{\langle [1], [2] \rangle, \langle [2], [1] \rangle\} \text{ and } \{\langle [1], [2] \rangle, \langle [2], [1] \rangle, \langle [2], [2] \rangle\}.$$

In either case, p and $\Box p \rightarrow p$ are false and $\Box p$ is true at $[1]$; p and $\Box p \rightarrow p$ are true and $\Box p$ is false at $[2]$.

Problem 5.3. Consider the following model $\mathfrak{M} = \langle W, R, V \rangle$ where $W = \{0\sigma : \sigma \in \mathbb{B}^*\}$, the set of sequences of 0s and 1s starting with 0, with $R\sigma\sigma'$ iff $\sigma' = \sigma 0$ or $\sigma' = \sigma 1$, and $V(p) = \{\sigma 0 : \sigma \in \mathbb{B}^*\}$ and $V(q) = \{\sigma 1 : \sigma \in \mathbb{B}^* \setminus \{1\}\}$. Here's a picture:



We have $\mathfrak{M}, w \not\models \Box(p \vee q) \rightarrow (\Box p \vee \Box q)$ for every w .

Let Γ be the set of sub-formulas of $\Box(p \vee q) \rightarrow (\Box p \vee \Box q)$. What are W^* and V^* ? What is the accessibility relation of the finest filtration of \mathfrak{M} ? Of the coarsest?

5.5 Filtrations are Finite

We've defined filtrations for any set Γ that is closed under sub-formulas. mod:fil:fin:sec Nothing in the definition itself guarantees that filtrations are finite. In fact, when Γ is infinite (e.g., is the set of all formulas), it may well be infinite. However, if Γ is finite (e.g., when it is the set of sub-formulas of a given formula φ), so is any filtration through Γ .

Proposition 5.12. *If Γ is finite then any filtration \mathfrak{M}^* of a model \mathfrak{M} through Γ is also finite.* mod:fil:fin:prop:filt-are-finite

Proof. The size of W^* is the number of different classes $[w]$ under the equivalence relation \equiv . Any two worlds u, v in such class—that is, any u and v such that $u \equiv v$ —agree on all formulas φ in Γ , $\varphi \in \Gamma$ either φ is true at both u and v , or at neither. So each class $[w]$ corresponds to subset of Γ , namely the set of all $\varphi \in \Gamma$ such that φ is true at the worlds in $[w]$. No two different classes $[u]$ and $[v]$ correspond to the same subset of Γ . For if the set of formulas true at u and that of formulas true at v are the same, then u and v agree on all formulas in Γ , i.e., $u \equiv v$. But then $[u] = [v]$. So, there is an injective function from W^* to $\wp(\Gamma)$, and hence $|W^*| \leq |\wp(\Gamma)|$. Hence if Γ contains n sentences, the cardinality of W^* is no greater than 2^n . \square

5.6 K and S5 have the Finite Model Property

mod:fil:fmp:
sec

Definition 5.13. A system Σ of modal logic is said to have the *finite model property* if whenever a formula φ is true at a world in a model of Σ then φ is true at a world in a *finite* model of Σ .

mod:fil:fmp:
prop:K-fmp

Proposition 5.14. **K** has the finite model property.

Proof. **K** is the set of valid formulas, i.e., any model is a model of **K**. By Theorem 5.5, if $\mathfrak{M}\varphi[w]$, then $\mathfrak{M}^*\varphi[w]$ for any filtration of \mathfrak{M} through the set Γ of sub-formulas of φ . Any formula only has finitely many sub-formulas, so Γ is finite. By Proposition 5.12, $|W^*| \leq 2^n$, where n is the number of formulas in Γ . And since **K** imposes no restriction on models, \mathfrak{M}^* is a **K**-model. \square

To show that a logic **L** has the finite model property via filtrations it is essential that the filtration of an **L**-model is itself a **L**-model. Often this requires a fair bit of work, and not any filtration yields a **L**-model. However, for universal models, this still holds.

mod:fil:fmp:
prop:univ-fin

Proposition 5.15. Let \mathcal{U} be the class of universal models (see Proposition 2.14) and \mathcal{U}_{Fin} the class of all finite universal models. Then any formula φ is valid in \mathcal{U} if and only if it is valid in \mathcal{U}_{Fin} .

Proof. Finite universal models are universal models, so the left-to-right direction is trivial. For the right-to-left direction, suppose that φ is false at some world w in a universal model \mathfrak{M} . Let Γ contain φ as well as all of its sub-formulas; clearly Γ is finite. Take a filtration \mathfrak{M}^* of \mathfrak{M} ; then \mathfrak{M}^* is finite by Proposition 5.12, and by Theorem 5.5, φ is false at $[w]$ in \mathfrak{M}^* . It remains to observe that \mathfrak{M}^* is also universal: given u and v , by hypothesis Ruv and by Definition 5.4(2), also $R^*[u][v]$. \square

mod:fil:fmp:
cor:S5fmp

Corollary 5.16. **S5** has the finite model property.

Proof. By Proposition 2.14, if φ is true at a world in some reflexive and euclidean model then it is true at a world in a universal model. By Proposition 5.15, it is true at a world in a finite universal model (namely the filtration of the model through the set of sub-formulas of φ). Every universal model is also reflexive and euclidean; so φ is true at a world in a finite reflexive euclidean model. \square

Problem 5.4. Show that any filtration of a serial or reflexive model is also serial or reflexive (respectively).

Problem 5.5. Find a non-symmetric (non-transitive, non-euclidean) filtration of a symmetric (transitive, euclidean) model.

5.7 S5 is Decidable

The finite model property gives us an easy way to show that systems of modal logic given by schemas are *decidable* (i.e., that there is a computable procedure to determine whether a formulas is *derivable* in the system or not). mod:fil:dec:
sec

Theorem 5.17. *S5 is decidable.*

Proof. Let φ be given, and suppose the propositional variables occurring in φ are among p_1, \dots, p_k . Since for each n there are only finitely many models with n worlds assigning a value to p_1, \dots, p_k , we can enumerate, *in parallel*, all the theorems of **S5** by generating proofs in some systematic way; and all the models containing 1, 2, \dots worlds and checking whether φ fails at a world in some such model. Eventually one of the two parallel processes will give an answer, as by [Theorem 4.17](#) and [Corollary 5.16](#), either φ is *derivable* or it fails in a finite universal model. \square

The above proof works for **S5** because filtrations of universal models are automatically universal. The same holds for reflexivity and seriality, but more work is needed for other properties.

Problem 5.6. Show that any filtration of a serial or reflexive model is also serial or reflexive (respectively).

Problem 5.7. Find a non-symmetric (non-transitive, non-euclidean) filtration of a symmetric (transitive, euclidean) model.

5.8 Filtrations and Properties of Accessibility

As noted, filtrations of universal, serial, and reflexive models are always also universal, serial, or reflexive. But not every filtration of a symmetric or transitive model is symmetric or transitive, respectively. In some cases, however, it is possible to define filtrations so that this does hold. In order to do so, we proceed as in the definition of the coarsest filtration, but add additional conditions to the definition of R^* . Let Γ be closed under sub-formulas. Consider the relations $C_i(u, v)$ in [table 5.1](#) between worlds u, v in a model $\mathfrak{M} = \langle W, R, V \rangle$. We can define $R^*[u][v]$ on the basis of combinations of these conditions. For instance, if we stipulate that $R^*[u][v]$ iff the condition $C_1(u, v)$ holds, we get exactly the coarsest filtration. If we stipulate $R^*[u][v]$ iff both $C_1(u, v)$ and $C_2(u, v)$ hold, we get a different filtration. It is “finer” than the coarsest since fewer pairs of worlds satisfy $C_1(u, v)$ and $C_2(u, v)$ than $C_1(u, v)$ alone. mod:fil:acc:
sec

Theorem 5.18. *Let $\mathfrak{M} = \langle W, R, P \rangle$ be a model, Γ closed under sub-formulas. Let W^* and V^* be defined as in [Definition 5.4](#). Then:* mod:fil:acc:
thm:more-filtrations

1. *Suppose $R^*[u][v]$ if and only if $C_1(u, v) \wedge C_2(u, v)$. Then R^* is symmetric, and $\mathfrak{M}^* = \langle W^*, R^*, V^* \rangle$ is a filtration if \mathfrak{M} is symmetric.*

$C_1(u, v)$:	if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Box\varphi$ then $\mathfrak{M}, v \Vdash \varphi$; and if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \varphi$ then $\mathfrak{M}, u \Vdash \Diamond\varphi$;
$C_2(u, v)$:	if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \Box\varphi$ then $\mathfrak{M}, u \Vdash \varphi$; and if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \varphi$ then $\mathfrak{M}, v \Vdash \Diamond\varphi$;
$C_3(u, v)$:	if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Box\varphi$ then $\mathfrak{M}, v \Vdash \Box\varphi$; and if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \Diamond\varphi$ then $\mathfrak{M}, u \Vdash \Diamond\varphi$;
$C_4(u, v)$:	if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \Box\varphi$ then $\mathfrak{M}, u \Vdash \Box\varphi$; and if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Diamond\varphi$ then $\mathfrak{M}, v \Vdash \Diamond\varphi$;

Table 5.1: Conditions on possible worlds for defining filtrations.

mod:fil:acc:
tab:Cn-filtrations

2. Suppose $R^*[u][v]$ if and only if $C_1(u, v) \wedge C_3(u, v)$. Then R^* is transitive, and $\mathfrak{M}^* = \langle W^*, R^*, V^* \rangle$ is a filtration if \mathfrak{M} is transitive.
3. Suppose $R^*[u][v]$ if and only if $C_1(u, v) \wedge C_2(u, v) \wedge C_3(u, v) \wedge C_4(u, v)$. Then R^* is symmetric and transitive, and $\mathfrak{M}^* = \langle W^*, R^*, V^* \rangle$ is a filtration if \mathfrak{M} is symmetric and transitive.
4. Suppose R^* is defined as $R^*[u][v]$ if and only if $C_1(u, v) \wedge C_3(u, v) \wedge C_4(u, v)$. Then R^* is transitive and euclidean, and $\mathfrak{M}^* = \langle W^*, R^*, V^* \rangle$ is a filtration if \mathfrak{M} is transitive and euclidean.

Proof. 1. It's immediate that R^* is symmetric, since $C_1(u, v) \Leftrightarrow C_2(v, u)$ and $C_2(u, v) \Leftrightarrow C_1(v, u)$. So it's left to show that if \mathfrak{M} is symmetric then \mathfrak{M}^* is a filtration through Γ . Condition $C_1(u, v)$ guarantees that (2b) and (2c) of Definition 5.4 are satisfied. So we just have to verify Definition 5.4(2a), i.e., that Ruv implies $R^*[u][v]$.

So suppose Ruv . To show $R^*[u][v]$ we need to establish that $C_1(u, v)$ and $C_2(u, v)$. For C_1 : if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \Box\varphi$ then also $\mathfrak{M}, v \Vdash \varphi$ (since Ruv). Similarly, if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \varphi$ then $\mathfrak{M}, u \Vdash \Diamond\varphi$ since Ruv . For C_2 : if $\Box\varphi \in \Gamma$ and $\mathfrak{M}, v \Vdash \Box\varphi$ then Ruv implies Rvu by symmetry, so that $\mathfrak{M}, u \Vdash \varphi$. Similarly, if $\Diamond\varphi \in \Gamma$ and $\mathfrak{M}, u \Vdash \varphi$ then $\mathfrak{M}, v \Vdash \Diamond\varphi$ (since Rvu by symmetry).

2. Exercise.
3. Exercise.
4. Exercise.

□

Problem 5.8. Complete the proof of Theorem 5.18.

5.9 Filtrations of Euclidean Models

mod:fil:euc:
sec

The approach of section 5.8 does not work in the case of models that are euclidean or serial and euclidean. Consider the model at the top of Figure 5.2,

which is both euclidean and serial. Let $\Gamma = \{p, \Box p\}$. When taking a filtration through Γ , then $[w_1] = [w_3]$ since w_1 and w_3 are the only worlds that agree on Γ . Any filtration will also have the arrow inherited from \mathfrak{M} , as depicted in Figure 5.3. That model isn't euclidean. Moreover, we cannot add arrows to that model in order to make it euclidean. We would have to add double arrows between $[w_2]$ and $[w_4]$, and then also between w_2 and w_5 . But $\Box p$ is supposed to be true at w_2 , while p is false at w_5 .

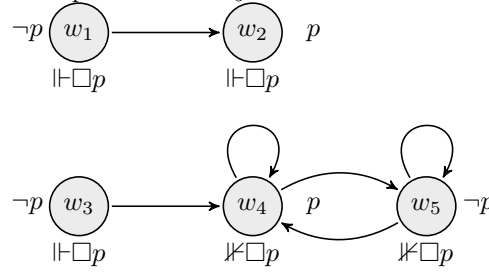


Figure 5.2: A serial and euclidean model.

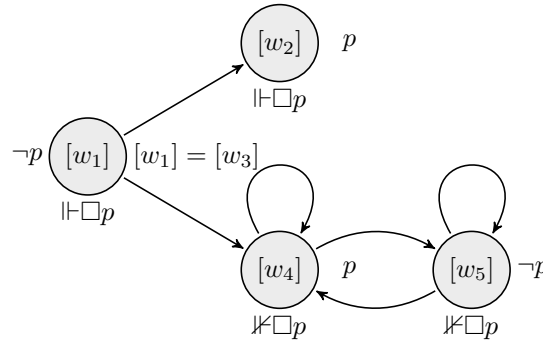


Figure 5.3: The filtration of the model in Figure 5.2.

In particular, to obtain a euclidean filtration it is not enough to consider filtrations through arbitrary Γ 's closed under sub-formulas. Instead we need to consider sets Γ that are *modally closed* (see Definition 5.1). Such sets of sentences are infinite, and therefore do not immediately yield a finite model property or the decidability of the corresponding system.

Theorem 5.19. *Let Γ be modally closed, $\mathfrak{M} = \langle W, R, V \rangle$, and $\mathfrak{M}^* = \langle W^*, R^*, V_m^* \rangle$ be a coarsest filtration of \mathfrak{M} .*

1. If \mathfrak{M} is symmetric, so is \mathfrak{M}^* .
2. If \mathfrak{M} is transitive, so is \mathfrak{M}^* .
3. If \mathfrak{M} is euclidean, so is \mathfrak{M}^* .

Proof. 1. If \mathfrak{M}^* is a coarsest filtration, then by definition $R^*[u][v]$ holds if and only if $C_1(u, v)$. For transitivity, suppose $C_1(u, v)$ and $C_1(v, w)$; we have to show $C_1(u, w)$. Suppose $\mathfrak{M}, u \Vdash \Box \varphi$; then $\mathfrak{M}, u \Vdash \Box \Box \varphi$

since 4 is valid in all transitive models; since $\Box\Box\varphi \in \Gamma$ by closure, also by $C_1(u, v)$, $\mathfrak{M}, v \Vdash \Box\varphi$ and by $C_1(v, w)$, also $\mathfrak{M}, w \Vdash \varphi$. Suppose $\mathfrak{M}, w \Vdash \varphi$; then $\mathfrak{M}, v \Vdash \Diamond\varphi$ by $C_1(v, w)$, since $\Diamond\varphi \in \Gamma$ by modal closure. By $C_1(u, v)$, we get $\mathfrak{M}, u \Vdash \Diamond\Diamond\varphi$ since $\Diamond\Diamond\varphi \in \Gamma$ by modal closure. Since 4_\Diamond is valid in all transitive models, $\mathfrak{M}, u \Vdash \Diamond\varphi$.

2. Exercise. Use the fact that both 5 and 5_\Diamond are valid in all euclidean models.
3. Exercise. Use the fact that B and B_\Diamond are valid in all symmetric models. □

Problem 5.9. Complete the proof of [Theorem 5.19](#).

Chapter 6

Modal Tableaux

Draft chapter on prefixed tableaux for modal logic. Needs more examples, completeness proofs, and discussion of how one can find countermodels from unsuccessful searches for closed tableaux.

6.1 Introduction

Tableaux are certain (downward-branching) trees of **signed formulas**, i.e., pairs consisting of a truth value sign (\mathbb{T} or \mathbb{F}) and a **sentence** mod:tab:int:sec

$$\mathbb{T}\varphi \text{ or } \mathbb{F}\varphi.$$

A **tableau** begins with a number of *assumptions*. Each further **signed formula** is generated by applying one of the inference rules. Some inference rules add one or more **signed formulas** to a tip of the tree; others add two new tips, resulting in two branches. Rules result in **signed formulas** where the **formula** is less complex than that of the **signed formula** to which it was applied. When a branch contains both $\mathbb{T}\varphi$ and $\mathbb{F}\varphi$, we say the branch is *closed*. If every branch in a **tableau** is closed, the entire **tableau** is closed. A closed **tableau** constitutes a **derivation** that shows that the set of **signed formulas** which were used to begin the **tableau** are unsatisfiable. This can be used to define a \vdash relation: $\Gamma \vdash \varphi$ iff there is some finite set $\Gamma_0 = \{\psi_1, \dots, \psi_n\} \subseteq \Gamma$ such that there is a closed **tableau** for the assumptions

$$\{\mathbb{F}\varphi, \mathbb{T}\psi_1, \dots, \mathbb{T}\psi_n\}.$$

For modal logics, we have to both extend the notion of **signed formula** and add rules that cover \Box and \Diamond . In addition to a sign (\mathbb{T} or \mathbb{F}), **formulas** in modal **tableaux** also have *prefixes* σ . The prefixes are non-empty sequences of positive integers, i.e., $\sigma \in (\mathbb{Z}^+)^* \setminus \{\Lambda\}$. When we write such prefixes without the surrounding $\langle \rangle$, and separate the individual **elements** by \cdot 's instead of $,$'s.

$\frac{\sigma \mathbb{T} \neg \varphi}{\sigma \mathbb{F} \varphi} \neg \mathbb{T}$	$\frac{\sigma \mathbb{F} \neg \varphi}{\sigma \mathbb{T} \varphi} \neg \mathbb{F}$
$\frac{\sigma \mathbb{T} \varphi \wedge \psi}{\sigma \mathbb{T} \varphi \quad \sigma \mathbb{T} \psi} \wedge \mathbb{T}$	$\frac{\sigma \mathbb{F} \varphi \wedge \psi}{\sigma \mathbb{F} \varphi \quad \quad \sigma \mathbb{F} \psi} \wedge \mathbb{F}$
$\frac{\sigma \mathbb{T} \varphi \vee \psi}{\sigma \mathbb{T} \varphi \quad \quad \sigma \mathbb{T} \psi} \vee \mathbb{T}$	$\frac{\sigma \mathbb{F} \varphi \vee \psi}{\sigma \mathbb{F} \varphi \quad \sigma \mathbb{F} \psi} \vee \mathbb{F}$
$\frac{\sigma \mathbb{T} \varphi \rightarrow \psi}{\sigma \mathbb{F} \varphi \quad \quad \sigma \mathbb{T} \psi} \rightarrow \mathbb{T}$	$\frac{\sigma \mathbb{F} \varphi \rightarrow \psi}{\sigma \mathbb{T} \varphi \quad \sigma \mathbb{F} \psi} \rightarrow \mathbb{F}$

mod:tab:rul:
tab:prop-rules

Table 6.1: Prefixed **tableau** rules for the propositional connectives

If σ is a prefix, then $\sigma.n$ is $\sigma \smallfrown \langle n \rangle$; e.g., if $\sigma = 1.2.1$, then $\sigma.3$ is $1.2.1.3$. So for instance,

$$1.2 \mathbb{T} \Box \varphi \rightarrow \varphi$$

is a *prefixed signed formula* (or just a *prefixed formula* for short).

Intuitively, the prefix names a world in a model that might satisfy the **formulas** on a branch of a **tableau**, and if σ names some world, then $\sigma.n$ names a world accessible from (the world named by) σ .

6.2 Rules for K

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sec

The rules for the regular propositional connectives are the same as for regular propositional signed **tableaux**, just with prefixes added. In each case, the rule applied to a signed **formula** $\sigma S \varphi$ produces new **formulas** that are also prefixed by σ . This should be intuitively clear: e.g., if $\varphi \wedge \psi$ is true at (a world named by) σ , then φ and ψ are true at σ (and not at any other world). We collect the propositional rules in [section 6.2](#).

The closure condition is the same as for ordinary **tableaux**, although we require that not just the **formulas** but also the prefixes must match. So a branch is closed if it contains both

$$\sigma \mathbb{T} \varphi \quad \text{and} \quad \sigma \mathbb{F} \varphi$$

for some prefix σ and **formula** φ .

The rules for setting up assumptions is also as for ordinary **tableaux**, except that for assumptions we always use the prefix 1. (It does not matter which

$\frac{\sigma \mathbb{T} \Box \varphi}{\sigma.n \mathbb{T} \varphi} \Box \mathbb{T}$ <p>$\sigma.n$ is used</p>	$\frac{\sigma \mathbb{F} \Box \varphi}{\sigma.n \mathbb{F} \varphi} \Box \mathbb{F}$ <p>$\sigma.n$ is new</p>
$\frac{\sigma \mathbb{T} \Diamond \varphi}{\sigma.n \mathbb{T} \varphi} \Diamond \mathbb{T}$ <p>$\sigma.n$ is new</p>	$\frac{\sigma \mathbb{F} \Diamond \varphi}{\sigma.n \mathbb{F} \varphi} \Diamond \mathbb{F}$ <p>$\sigma.n$ is used</p>

Table 6.2: The modal rules for K.

mod:tab:rul:
tab:rules-K

prefix we use, as long as it's the same for all assumptions.) So, e.g., we say that

$$\psi_1, \dots, \psi_n \vdash \varphi$$

iff there is a closed tableau for the assumptions

$$1 \mathbb{T} \psi_1, \dots, 1 \mathbb{T} \psi_n, 1 \mathbb{F} \varphi.$$

For the modal operators \Box and \Diamond , the prefix of the conclusion of the rule applied to a formula with prefix σ is $\sigma.n$. However, which n is allowed depends on whether the sign is \mathbb{T} or \mathbb{F} .

The $\mathbb{T}\Box$ rule extends a branch containing $\sigma \mathbb{T}\Box\varphi$ by $\sigma.n \mathbb{T}\varphi$. Similarly, the $\mathbb{F}\Diamond$ rule extends a branch containing $\sigma \mathbb{F}\Diamond\varphi$ by $\sigma.n \mathbb{F}\varphi$. They can only be applied for a prefix $\sigma.n$ which *already* occurs on the branch in which it is applied. Let's call such a prefix "used" (on the branch).

The $\mathbb{F}\Box$ rule extends a branch containing $\sigma \mathbb{F}\Box\varphi$ by $\sigma.n \mathbb{F}\varphi$. Similarly, the $\mathbb{T}\Diamond$ rule extends a branch containing $\sigma \mathbb{T}\Diamond\varphi$ by $\sigma.n \mathbb{T}\varphi$. These rules, however, can only be applied for a prefix $\sigma.n$ which *does not* already occur on the branch in which it is applied. We call such prefixes "new" (to the branch).

The rules are given in table 6.2.

The requirements that the restriction that the prefix for $\Box\mathbb{T}$ must be used is necessary as otherwise we would count the following as a closed tableau:

1. $1 \mathbb{T} \Box \varphi$ Assumption
 2. $1 \mathbb{F} \Diamond \varphi$ Assumption
 3. $1.1 \mathbb{T} \varphi$ $\Box\mathbb{T} 1$
 4. $1.1 \mathbb{F} \varphi$ $\Diamond\mathbb{F} 2$
- \otimes

But $\Box\varphi \neq \Diamond\varphi$, so our proof system would be unsound. Likewise, $\Diamond\varphi \neq \Box\varphi$, but without the restriction that the prefix for $\Box\mathbb{F}$ must be new, this would be a closed tableau:

1.	1 T	$\Diamond\varphi$	Assumption
2.	1 F	$\Box\varphi$	Assumption
3.	1.1 T	φ	$\Diamond\mathbb{T}$ 1
4.	1.1 F	φ	$\Box\mathbb{F}$ 2
		\otimes	

6.3 Tableaux for \mathbf{K}

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sec

Example 6.1. We give a closed tableau that shows $\vdash (\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$.

1.	1 F	$(\Box\varphi \wedge \Box\psi) \rightarrow \Box(\varphi \wedge \psi)$	Assumption
2.	1 T	$\Box\varphi \wedge \Box\psi$	$\rightarrow\mathbb{T}$ 1
3.	1 F	$\Box(\varphi \wedge \psi)$	$\rightarrow\mathbb{T}$ 1
4.	1 T	$\Box\varphi$	$\wedge\mathbb{T}$ 2
5.	1 T	$\Box\psi$	$\wedge\mathbb{T}$ 2
6.	1.1 F	$\varphi \wedge \psi$	$\Box\mathbb{F}$ 3
		\swarrow	
7.	1.1 F	φ	$\wedge\mathbb{F}$ 6
		1.1 F	ψ
8.	1.1 T	φ	$\Box\mathbb{T}$ 4; $\Box\mathbb{T}$ 5
		1.1 T	ψ
		\otimes	

Example 6.2. We give a closed tableau that shows $\vdash \Diamond(\varphi \vee \psi) \rightarrow (\Diamond\varphi \vee \Diamond\psi)$:

1.	1 F	$\Diamond(\varphi \vee \psi) \rightarrow (\Diamond\varphi \vee \Diamond\psi)$	Assumption
2.	1 T	$\Diamond(\varphi \vee \psi)$	$\rightarrow\mathbb{T}$ 1
3.	1 F	$\Diamond\varphi \vee \Diamond\psi$	$\rightarrow\mathbb{T}$ 1
4.	1 F	$\Diamond\varphi$	$\vee\mathbb{F}$ 3
5.	1 F	$\Diamond\psi$	$\vee\mathbb{F}$ 3
6.	1.1 T	$\varphi \vee \psi$	$\Diamond\mathbb{T}$ 2
		\swarrow	
7.	1.1 T	φ	$\vee\mathbb{T}$ 6
		1.1 T	ψ
8.	1.1 F	φ	$\Diamond\mathbb{F}$ 4; $\Diamond\mathbb{F}$ 5
		1.1 F	ψ
		\otimes	

Problem 6.1. Find closed tableaux in \mathbf{K} for the following formulas:

1. $\Box\neg p \rightarrow \Box(p \rightarrow q)$
2. $(\Box p \vee \Box q) \rightarrow \Box(p \vee q)$
3. $\Diamond p \rightarrow \Diamond(p \vee q)$

6.4 Soundness

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sec

This soundness proof reuses the soundness proof for classical propositional logic, i.e., it proves everything from scratch. That's ok if you want a self-contained soundness proof. If you already have seen soundness for ordinary tableau this will be repetitive. It's planned to make it possible to switch between self-contained version and a version building on the non-modal case.

[explanation](#) In order to show that prefixed **tableau** are sound, we have to show that if

$$1\mathbb{T}\psi_1, \dots, 1\mathbb{T}\psi_n, 1\mathbb{F}\varphi$$

has a closed **tableau** then $\psi_1, \dots, \psi_n \models \varphi$. It is easier to prove the contrapositive: if for some \mathfrak{M} and world w , $\mathfrak{M}, w \Vdash \psi_i$ for all $i = 1, \dots, n$ but $\mathfrak{M}, w \not\Vdash \varphi$, then no **tableau** can close. Such a countermodel shows that the initial assumptions of the **tableau** are satisfiable. The strategy of the proof is to show that whenever all the prefixed **formulas** on a **tableau** branch are satisfiable, any application of a rule results in at least one extended branch that is also satisfiable. Since closed branches are unsatisfiable, any **tableau** for a satisfiable set of prefixed **formulas** must have at least one open branch.

In order to apply this strategy in the modal case, we have to extend our definition of “satisfiable” to modal modals and prefixes. With that in hand, however, the proof is straightforward.

Definition 6.3. Let P be some set of prefixes, i.e., $P \subseteq (\mathbb{Z}^+)^* \setminus \{A\}$ and let \mathfrak{M} be a model. A function $f: P \rightarrow W$ is an *interpretation of P* in \mathfrak{M} if, whenever σ and $\sigma.n$ are both in P , then $Rf(\sigma)f(\sigma.n)$.

Relative to an interpretation of prefixes P we can define:

1. \mathfrak{M} satisfies $\sigma\mathbb{T}\varphi$ iff $\mathfrak{M}, f(\sigma) \Vdash \varphi$.
2. \mathfrak{M} satisfies $\sigma\mathbb{F}\varphi$ iff $\mathfrak{M}, f(\sigma) \not\Vdash \varphi$.

Definition 6.4. Let Γ be a set of prefixed **formulas**, and let $P(\Gamma)$ be the set of prefixes that occur in it. If f is an interpretation of $P(\Gamma)$ in \mathfrak{M} , we say that \mathfrak{M} satisfies Γ with respect to f , $\mathfrak{M}, f \Vdash \Gamma$, if \mathfrak{M} satisfies every prefixed **formula** in Γ with respect to f . Γ is *satisfiable* iff there is a model \mathfrak{M} and interpretation f of $P(\Gamma)$ such that $\mathfrak{M}, f \Vdash \Gamma$.

Proposition 6.5. *If Γ contains both $\sigma\mathbb{T}\varphi$ and $\sigma\mathbb{F}\varphi$, for some **formula** φ and prefix σ , then Γ is unsatisfiable.*

Proof. There cannot be a model \mathfrak{M} and interpretation f of $P(\Gamma)$ such that both $\mathfrak{M}, f(\sigma) \Vdash \varphi$ and $\mathfrak{M}, f(\sigma) \not\Vdash \varphi$. \square

Theorem 6.6 (Soundness). *If Γ has a closed tableau, Γ is unsatisfiable.*

Proof. We call a branch of a tableau satisfiable iff the set of signed formulas on it is satisfiable, and let's call a tableau satisfiable if it contains at least one satisfiable branch.

We show the following: Extending a satisfiable tableau by one of the rules of inference always results in a satisfiable tableau. This will prove the theorem: any closed tableau results by applying rules of inference to the tableau consisting only of assumptions from Γ . So if Γ were satisfiable, any tableau for it would be satisfiable. A closed tableau, however, is clearly not satisfiable, since all its branches are closed and closed branches are unsatisfiable.

Suppose we have a satisfiable tableau, i.e., a tableau with at least one satisfiable branch. Applying a rule of inference either adds signed formulas to a branch, or splits a branch in two. If the tableau has a satisfiable branch which is not extended by the rule application in question, it remains a satisfiable branch in the extended tableau, so the extended tableau is satisfiable. So we only have to consider the case where a rule is applied to a satisfiable branch.

Let Γ be the set of signed formulas on that branch, and let $\sigma S\varphi \in \Gamma$ be the signed formula to which the rule is applied. If the rule does not result in a split branch, we have to show that the extended branch, i.e., Γ together with the conclusions of the rule, is still satisfiable. If the rule results in split branch, we have to show that at least one of the two resulting branches is satisfiable. First, we consider the possible inferences with only one premise.

1. The branch is expanded by applying $\neg\mathbb{T}$ to $\sigma\mathbb{T}\neg\psi \in \Gamma$. Then the extended branch contains the signed formulas $\Gamma \cup \{\sigma\mathbb{F}\psi\}$. Suppose $\mathfrak{M}, f \Vdash \Gamma$. In particular, $\mathfrak{M}, f(\sigma) \Vdash \neg\psi$. Thus, $\mathfrak{M}, f(\sigma) \not\Vdash \psi$, i.e., \mathfrak{M} satisfies $\sigma\mathbb{F}\psi$ with respect to f .
2. The branch is expanded by applying $\neg\mathbb{F}$ to $\sigma\mathbb{F}\neg\psi \in \Gamma$: Exercise.
3. The branch is expanded by applying $\wedge\mathbb{T}$ to $\sigma\mathbb{T}\psi \wedge \chi \in \Gamma$, which results in two new signed formulas on the branch: $\sigma\mathbb{T}\psi$ and $\sigma\mathbb{T}\chi$. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular $\mathfrak{M}, f(\sigma) \Vdash \psi \wedge \chi$. Then $\mathfrak{M}, f(\sigma) \Vdash \psi$ and $\mathfrak{M}, f(\sigma) \Vdash \chi$. This means that \mathfrak{M} satisfies both $\sigma\mathbb{T}\psi$ and $\sigma\mathbb{T}\chi$ with respect to f .
4. The branch is expanded by applying $\vee\mathbb{F}$ to $\mathbb{T}\psi \vee \chi \in \Gamma$: Exercise.
5. The branch is expanded by applying $\rightarrow\mathbb{F}$ to $\sigma\mathbb{F}\psi \rightarrow \chi \in \Gamma$: This results in two new signed formulas on the branch: $\sigma\mathbb{T}\psi$ and $\sigma\mathbb{F}\chi$. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular $\mathfrak{M}, f(\sigma) \not\Vdash \psi \rightarrow \chi$. Then $\mathfrak{M}, f(\sigma) \Vdash \psi$ and $\mathfrak{M}, f(\sigma) \not\Vdash \chi$. This means that \mathfrak{M}, f satisfies both $\sigma\mathbb{T}\psi$ and $\sigma\mathbb{F}\chi$.
6. The branch is expanded by applying $\Box\mathbb{T}$ to $\sigma\mathbb{T}\Box\psi \in \Gamma$: This results in a new signed formula $\sigma.n\mathbb{T}\psi$ on the branch, for some $\sigma.n \in P(\Gamma)$ (since $\sigma.n$ must be used). Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$. Since f is an interpretation of prefixes and both $\sigma, \sigma.n \in P(\Gamma)$, we know that $Rf(\sigma)f(\sigma.n)$. Hence, $\mathfrak{M}, f(\sigma.n) \Vdash \psi$, i.e., \mathfrak{M}, f satisfies $\sigma.n\mathbb{T}\psi$.

7. The branch is expanded by applying $\Box\mathbb{F}$ to $\sigma\mathbb{F}\Box\psi \in \Gamma$: This results in a new **signed formula** $\sigma.n\mathbb{F}\varphi$, where $\sigma.n$ is a new prefix on the branch, i.e., $\sigma.n \notin P(\Gamma)$. Since Γ is satisfiable, there is a \mathfrak{M} and interpretation f of $P(\Gamma)$ such that $\mathfrak{M}, f \models \Gamma$, in particular $\mathfrak{M}, f(\sigma) \not\models \Box\psi$. We have to show that $\Gamma \cup \{\sigma.n\mathbb{F}\psi\}$ is satisfiable. To do this, we define an interpretation of $P(\Gamma) \cup \{\sigma.n\}$ as follows:

Since $\mathfrak{M}, f(\sigma) \not\models \Box\psi$, there is a $w \in W$ such that $Rf(\sigma)w$ and $\mathfrak{M}, w \not\models \psi$. Let f' be like f , except that $f'(\sigma.n) = w$. Since $f'(\sigma) = f(\sigma)$ and $Rf(\sigma)w$, we have $Rf'(\sigma)f'(\sigma.n)$, so f' is an interpretation of $P(\Gamma) \cup \{\sigma.n\}$. Obviously $\mathfrak{M}, f'(\sigma.n) \not\models \psi$. Since $f(\sigma') = f'(\sigma')$ for all prefixes $\sigma' \in P(\Gamma)$, $\mathfrak{M}, f' \models \Gamma$. So, \mathfrak{M}, f' satisfies $\Gamma \cup \{\sigma.n\mathbb{F}\psi\}$.

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

1. The branch is expanded by applying $\wedge\mathbb{F}$ to $\sigma\mathbb{F}\psi \wedge \chi \in \Gamma$, which results in two branches, a left one continuing through $\sigma\mathbb{F}\psi$ and a right one through $\sigma\mathbb{F}\chi$. Suppose $\mathfrak{M}, f \models \Gamma$, in particular $\mathfrak{M}, f(\sigma) \not\models \psi \wedge \chi$. Then $\mathfrak{M}, f(\sigma) \not\models \psi$ or $\mathfrak{M}, f(\sigma) \not\models \chi$. In the former case, \mathfrak{M}, f satisfies $\sigma\mathbb{F}\psi$, i.e., the left branch is satisfiable. In the latter, \mathfrak{M}, f satisfies $\sigma\mathbb{F}\chi$, i.e., the right branch is satisfiable.
2. The branch is expanded by applying $\vee\mathbb{T}$ to $\mathbb{T}\psi \vee \chi \in \Gamma$: Exercise.
3. The branch is expanded by applying $\rightarrow\mathbb{T}$ to $\mathbb{T}\psi \rightarrow \chi \in \Gamma$: Exercise.

□

Problem 6.2. Complete the proof of [Theorem 6.6](#).

Corollary 6.7. *If $\Gamma \vdash \varphi$ then $\Gamma \models \varphi$.*

*mod:tab:sou:
cor:entailment-soundness*

Proof. If $\Gamma \vdash \varphi$ then for some $\psi_1, \dots, \psi_n \in \Gamma$, $\Delta = \{1\mathbb{F}\varphi, 1\mathbb{T}\psi_1, \dots, 1\mathbb{T}\psi_n\}$ has a closed **tableau**. We want to show that $\Gamma \models \varphi$. Suppose not, so for some \mathfrak{M} and w , $\mathfrak{M}, w \models \psi_i$ for $i = 1, \dots, n$, but $\mathfrak{M}, w \not\models \varphi$. Let $f(1) = w$; then f is an interpretation of $P(\Delta)$ into \mathfrak{M} , and \mathfrak{M} satisfies Δ with respect to f . But by [Theorem 6.6](#), Δ is unsatisfiable since it has a closed **tableau**, a contradiction. So we must have $\Gamma \vdash \varphi$ after all. □

Corollary 6.8. *If $\vdash \varphi$ then φ is true in all models.*

*mod:tab:sou:
cor:weak-soundness*

6.5 Rules for Other Accessibility Relations

In order to deal with logics determined by special accessibility relations, we consider the additional rules in [table 6.3](#).

*mod:tab:mru:
sec*

Adding these rules results in systems that are sound and complete for the logics given in [section 6.5](#).

$\frac{\sigma \mathbf{T} \Box \varphi}{\sigma \mathbf{T} \varphi} \mathbf{T} \Box$	$\frac{\sigma \mathbf{F} \Diamond \varphi}{\sigma \mathbf{F} \varphi} \mathbf{T} \Diamond$
$\frac{\sigma \mathbf{T} \Box \varphi}{\sigma \mathbf{T} \Diamond \varphi} \mathbf{D} \Box$	$\frac{\sigma \mathbf{F} \Diamond \varphi}{\sigma \mathbf{F} \Box \varphi} \mathbf{D} \Diamond$
$\frac{\sigma.n \mathbf{T} \Box \varphi}{\sigma \mathbf{T} \varphi} \mathbf{B} \Box$	$\frac{\sigma.n \mathbf{F} \Diamond \varphi}{\sigma \mathbf{F} \varphi} \mathbf{B} \Diamond$
$\frac{\sigma \mathbf{T} \Box \varphi}{\sigma.n \mathbf{T} \Box \varphi} 4 \Box$ $\sigma.n$ is used	$\frac{\sigma \mathbf{F} \Diamond \varphi}{\sigma.n \mathbf{F} \Diamond \varphi} 4 \Diamond$ $\sigma.n$ is used
$\frac{\sigma.n \mathbf{T} \Box \varphi}{\sigma \mathbf{T} \Box \varphi} 4r \Box$	$\frac{\sigma.n \mathbf{F} \Diamond \varphi}{\sigma \mathbf{F} \Diamond \varphi} 4r \Diamond$

Table 6.3: More modal rules.

mod:tab:mru:
tab:more-rules

Logic	R is ...	Rules
T = KT	reflexive	$\mathbf{T} \Box, \mathbf{T} \Diamond$
D = KD	serial	$\mathbf{D} \Box, \mathbf{D} \Diamond$
K4	transitive	$4 \Box, 4 \Diamond$
B = KTB	reflexive, symmetric	$\mathbf{T} \Box, \mathbf{T} \Diamond$ $\mathbf{B} \Box, \mathbf{B} \Diamond$
S4 = KT4	reflexive, transitive	$\mathbf{T} \Box, \mathbf{T} \Diamond,$ $4 \Box, 4 \Diamond$
S5 = KT4B	reflexive, transitive, euclidean	$\mathbf{T} \Box, \mathbf{T} \Diamond,$ $4 \Box, 4 \Diamond,$ $4r \Box, 4r \Diamond$

Table 6.4: **Tableau** rules for various modal logics.

mod:tab:mru:
tab:logics-rules

6.6 Tableaux for Other Logics

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sec

Example 6.9. We give a closed tableau that shows $\mathbf{S5} \vdash 5$, i.e., $\Box\varphi \rightarrow \Box\Diamond\varphi$.

1.	1F	$\Box\varphi \rightarrow \Box\Diamond\varphi$	Assumption
2.	1T	$\Box\varphi$	$\rightarrow\text{F } 1$
3.	1F	$\Box\Diamond\varphi$	$\rightarrow\text{F } 1$
4.	1.1F	$\Diamond\varphi$	$\Box\text{F } 3$
5.	1F	$\Diamond\varphi$	$4\text{r}\Diamond 4$
6.	1.1F	φ	$\Diamond\text{F } 5$
7.	1.1T	φ	$\Box\text{T } 2$
		\otimes	

Problem 6.3. Give closed tableaux that show the following:

1. $\mathbf{KT5} \vdash \text{B}$;
2. $\mathbf{KT5} \vdash 4$;
3. $\mathbf{KDB4} \vdash \text{T}$;
4. $\mathbf{KB4} \vdash 5$;
5. $\mathbf{KB5} \vdash 4$;
6. $\mathbf{KT} \vdash \text{D}$.

6.7 Soundness for Additional Rules

We say a rule is sound for a class of models if, whenever a branch in a tableau is satisfiable in a model from that class, the branch resulting from applying the rule is also satisfiable in a model from that class.

mod:tab:msn:
sec

Proposition 6.10. $\text{T}\Box$ and $\text{T}\Diamond$ are sound for reflexive models.

mod:tab:msn:
prop:soundness-T

Proof. 1. The branch is expanded by applying $\text{T}\Box$ to $\sigma\text{T}\Box\psi \in \Gamma$: This results in a new signed formula $\sigma\text{T}\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$. Since R is reflexive, we know that $Rf(\sigma)f(\sigma)$. Hence, $\mathfrak{M}, f(\sigma) \Vdash \psi$, i.e., \mathfrak{M}, f satisfies $\sigma\text{T}\psi$.

2. The branch is expanded by applying $\text{T}\Diamond$ to $\sigma\text{F}\Diamond\psi \in \Gamma$: This results in a new signed formula $\sigma\text{F}\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \not\Vdash \Diamond\psi$. Since R is reflexive, we know that $Rf(\sigma)f(\sigma)$. Hence, $\mathfrak{M}, f(\sigma) \not\Vdash \psi$, i.e., \mathfrak{M}, f satisfies $\sigma\text{F}\psi$. \square

Proposition 6.11. $\text{D}\Box$ and $\text{D}\Diamond$ are sound for serial models.

mod:tab:msn:
prop:soundness-D

- Proof.* 1. The branch is expanded by applying $D\Box$ to $\sigma\mathbb{T}\Box\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{T}\Diamond\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$. Since R is serial, there is a $w \in W$ such that $Rf(\sigma)w$. Then $\mathfrak{M}, w \Vdash \psi$, and hence $\mathfrak{M}, f(\sigma) \Vdash \Diamond\psi$. So, \mathfrak{M}, f satisfies $\sigma\mathbb{T}\Diamond\psi$.
2. The branch is expanded by applying $D\Diamond$ to $\sigma\mathbb{F}\Diamond\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{F}\Box\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \not\Vdash \Diamond\psi$. Since R is serial, there is a $w \in W$ such that $Rf(\sigma)w$. Then $\mathfrak{M}, w \not\Vdash \psi$, and hence $\mathfrak{M}, f(\sigma) \not\Vdash \Box\psi$. So, \mathfrak{M}, f satisfies $\sigma\mathbb{F}\Box\psi$. \square

mod:tab:msn:
prop:soundness-B

Proposition 6.12. $B\Box$ and $B\Diamond$ are sound for symmetric models.

- Proof.* 1. The branch is expanded by applying $B\Box$ to $\sigma.n\mathbb{T}\Box\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{T}\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma.n) \Vdash \Box\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} , we know that $Rf(\sigma)f(\sigma.n)$. Since R is symmetric, $Rf(\sigma.n)f(\sigma)$. Since $\mathfrak{M}, f(\sigma.n) \Vdash \Box\psi$, $\mathfrak{M}, f(\sigma) \Vdash \psi$. Hence, \mathfrak{M}, f satisfies $\sigma\mathbb{T}\psi$.
2. The branch is expanded by applying $B\Diamond$ to $\sigma.n\mathbb{F}\Diamond\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{F}\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma.n) \not\Vdash \Diamond\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} , we know that $Rf(\sigma)f(\sigma.n)$. Since R is symmetric, $Rf(\sigma.n)f(\sigma)$. Since $\mathfrak{M}, f(\sigma.n) \not\Vdash \Diamond\psi$, $\mathfrak{M}, f(\sigma) \not\Vdash \psi$. Hence, \mathfrak{M}, f satisfies $\sigma\mathbb{F}\psi$. \square

mod:tab:msn:
prop:soundness-4

Proposition 6.13. $4\Box$ and $4\Diamond$ are sound for transitive models.

- Proof.* 1. The branch is expanded by applying $4\Box$ to $\sigma\mathbb{T}\Box\psi \in \Gamma$: This results in a new **signed formula** $\sigma.n\mathbb{T}\Box\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} and $\sigma.n$ must be used, we know that $Rf(\sigma)f(\sigma.n)$. Now let w be any world such that $Rf(\sigma.n)w$. Since R is transitive, $Rf(\sigma)w$. Since $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$, $\mathfrak{M}, w \Vdash \psi$. Hence, $\mathfrak{M}, f(\sigma.n) \Vdash \Box\psi$, and \mathfrak{M}, f satisfies $\sigma.n\mathbb{T}\Box\psi$.
2. The branch is expanded by applying $4\Diamond$ to $\sigma\mathbb{F}\Diamond\psi \in \Gamma$: This results in a new **signed formula** $\sigma.n\mathbb{F}\Diamond\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma) \not\Vdash \Diamond\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} and $\sigma.n$ must be used, we know that $Rf(\sigma)f(\sigma.n)$. Now let w be any world such that $Rf(\sigma.n)w$. Since R is transitive, $Rf(\sigma)w$. Since $\mathfrak{M}, f(\sigma) \not\Vdash \Diamond\psi$, $\mathfrak{M}, w \not\Vdash \psi$. Hence, $\mathfrak{M}, f(\sigma.n) \not\Vdash \Diamond\psi$, and \mathfrak{M}, f satisfies $\sigma.n\mathbb{F}\Diamond\psi$. \square

mod:tab:msn:
prop:soundness-4r

Proposition 6.14. $4r\Box$ and $4r\Diamond$ are sound for euclidean models.

Proof. 1. The branch is expanded by applying $4r\Box$ to $\sigma.n\mathbb{T}\Box\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{T}\Box\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma.n) \Vdash \Box\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} , we know that $Rf(\sigma)f(\sigma.n)$. Now let w be any world such that $Rf(\sigma)w$. Since R is euclidean, $Rf(\sigma.n)w$. Since $\mathfrak{M}, f(\sigma.n) \Vdash \Box\psi$, $\mathfrak{M}, w \Vdash \psi$. Hence, $\mathfrak{M}, f(\sigma) \Vdash \Box\psi$, and \mathfrak{M}, f satisfies $\sigma\mathbb{T}\Box\psi$.

2. The branch is expanded by applying $4r\Diamond$ to $\sigma.n\mathbb{F}\Diamond\psi \in \Gamma$: This results in a new **signed formula** $\sigma\mathbb{T}\Box\psi$ on the branch. Suppose $\mathfrak{M}, f \Vdash \Gamma$, in particular, $\mathfrak{M}, f(\sigma.n) \not\Vdash \Diamond\psi$. Since f is an interpretation of prefixes on the branch into \mathfrak{M} , we know that $Rf(\sigma)f(\sigma.n)$. Now let w be any world such that $Rf(\sigma)w$. Since R is euclidean, $Rf(\sigma.n)w$. Since $\mathfrak{M}, f(\sigma.n) \not\Vdash \Diamond\psi$, $\mathfrak{M}, w \not\Vdash \psi$. Hence, $\mathfrak{M}, f(\sigma) \not\Vdash \Diamond\psi$, and \mathfrak{M}, f satisfies $\sigma\mathbb{F}\Diamond\psi$. \square

Corollary 6.15. *The **tableau** systems given in section 6.5 are sound for the respective classes of models.*

mod:tab:msn:
cor:soundness-logics

6.8 Simple Tableaux for S5

S5 is sound and complete with respect to the class of universal models, i.e., models where every world is accessible from every world. In universal models the accessibility relation doesn't matter: "there is a world w where $\mathfrak{M}, w \Vdash \varphi$ " is true if and only if there is such a w that's accessible from u . So in **S5**, we can define models as simply a set of worlds and a valuation V . This suggests that we should be able to simplify the **tableau** rules as well. In the general case, we take as prefixes sequences of positive integers, so that we can keep track of which such prefixes name worlds which are accessible from others: $\sigma.n$ names a world accessible from σ . But in **S5** any world is accessible from any world, so there is no need to so keep track. Instead, we can use positive integers as prefixes. The simplified rules are given in table 6.5.

mod:tabs:5:
sec

Example 6.16. We give a simplified closed tableau that shows $\mathbf{S5} \vdash 5$, i.e., $\Diamond\varphi \rightarrow \Box\Diamond\varphi$.

1.	$1\mathbb{F} \Diamond\varphi \rightarrow \Box\Diamond\varphi$	Assumption
2.	$1\mathbb{T} \Diamond\varphi$	$\rightarrow\mathbb{F} 1$
3.	$1\mathbb{F} \Box\Diamond\varphi$	$\rightarrow\mathbb{F} 1$
4.	$2\mathbb{F} \Diamond\varphi$	$\Box\mathbb{F} 3$
5.	$3\mathbb{T} \varphi$	$\Diamond\mathbb{T} 2$
6.	$3\mathbb{F} \varphi$	$\Diamond\mathbb{F} 4$
	\otimes	

6.9 Completeness for K

explanation To show that the method of **tableaux** is complete, we have to show that whenever there is no closed **tableau** to show $\Gamma \vdash \varphi$, then $\Gamma \not\Vdash \varphi$, i.e., there is

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sec

$\frac{n \mathbb{T} \Box \varphi}{m \mathbb{T} \varphi} \Box \mathbb{T}$ <p>m is used</p>	$\frac{n \mathbb{F} \Box \varphi}{m \mathbb{F} \varphi} \Box \mathbb{F}$ <p>m is new</p>
$\frac{n \mathbb{T} \Diamond \varphi}{m \mathbb{T} \varphi} \Diamond \mathbb{T}$ <p>m is new</p>	$\frac{n \mathbb{F} \Diamond \varphi}{m \mathbb{F} \varphi} \Diamond \mathbb{F}$ <p>m is used</p>

Table 6.5: Simplified rules for **S5**.

mod:tab:s5:
tab:rules-S5

a countermodel. But “there is no closed **tableau**” means that every way we could try to construct one has to fail to close. The trick is to see that if every such way fails to close, then a specific, *systematic and exhaustive* way also fails to close. And this systematic and exhaustive way would close if a closed **tableau** exists. The single tableau will contain, among its open branches, all the information required to define a countermodel. The countermodel given by an open branch in this tableau will contain all the prefixes used on that branch as the worlds, and a **propositional variable** p is true at σ iff $\sigma \mathbb{T} p$ occurs on the branch.

Definition 6.17. A branch in a **tableau** is called complete if, whenever it contains a prefixed **formula** $\sigma S \varphi$ to which a rule can be applied, it also contains

1. the prefixed **formulas** that are the corresponding conclusions of the rule, in the case of propositional stacking rules;
2. one of the corresponding conclusion **formulas** in the case of propositional branching rules;
3. at least one possible conclusion in the case of modal rules that require a new prefix;
4. the corresponding conclusion for every prefix occurring on the branch in the case of modal rules that require a used prefix.

For instance, a complete branch contains $\sigma \mathbb{T} \psi$ and $\sigma \mathbb{T} \chi$ whenever it contains $\mathbb{T} \psi \wedge \chi$. If it contains $\sigma \mathbb{T} \psi \vee \chi$ it contains at least one of $\sigma \mathbb{F} \psi$ and $\sigma \mathbb{T} \chi$. If it contains $\sigma \mathbb{F} \Box$ it also contains $\sigma.n \mathbb{F} \Box$ for at least one n . And whenever it contains $\sigma \mathbb{T} \Box$ it also contains $\sigma.n \mathbb{T} \Box$ for every n such that $\sigma.n$ is used on the branch.

explanation

Proposition 6.18. *Every finite Γ has a **tableau** in which every branch is complete.* mod:tab:cpl:
prop:complete-tableau

Proof. Consider an open branch in a **tableau** for Γ . There are finitely many prefixed **formulas** in the branch to which a rule could be applied. In some fixed order (say, top to bottom), for each of these prefixed **formulas** for which the conditions (1)–(4) do not already hold, apply the rules that can be applied to it to extend the branch. In some cases this will result in branching; apply the rule at the tip of each resulting branch for all remaining prefixed **formulas**. Since the number of prefixed **formulas** is finite, and the number of used prefixes on the branch is finite, this procedure eventually results in (possibly many) branches extending the original branch. Apply the procedure to each, and repeat. But by construction, every branch is closed. \square

Theorem 6.19 (Completeness). *If Γ has no closed **tableau**, Γ is satisfiable.* mod:tab:cpl:
thm:tableau-completeness

Proof. By the proposition, Γ has a **tableau** in which every branch is complete. Since it has no closed **tableau**, it has a **tableau** in which at least one branch is open and complete. Let Δ be the set of prefixed **formulas** on the branch, and $P(\Delta)$ the set of prefixes occurring in it.

We define a model $\mathfrak{M}(\Delta) = \langle P(\Delta), R, V \rangle$ where the worlds are the prefixes occurring in Δ , the accessibility relation is given by:

$$R\sigma\sigma' \quad \text{iff} \quad \sigma' = \sigma.n \quad \text{for some } n$$

and

$$V(p) = \{\sigma : \sigma \mathbb{T} p \in \Delta\}.$$

We show by induction on φ that if $\sigma \mathbb{T} \varphi \in \Delta$ then $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$, and if $\sigma \mathbb{F} \varphi \in \Delta$ then $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.

1. $\varphi \equiv p$: If $\sigma \mathbb{T} \varphi \in \Delta$ then $\sigma \in V(p)$ (by definition of V) and so $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
If $\sigma \mathbb{F} \varphi \in \Delta$ then $\sigma \mathbb{T} \varphi \notin \Delta$, since the branch would otherwise be closed. So $\sigma \notin V(p)$ and thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.
2. $\varphi \equiv \neg\psi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then $\sigma \mathbb{F} \psi \in \Delta$ since the branch is complete. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma \not\Vdash \psi$ and thus $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
If $\sigma \mathbb{F} \varphi \in \Delta$, then $\sigma \mathbb{T} \psi \in \Delta$ since the branch is complete. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma \Vdash \psi$ and thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.
3. $\varphi \equiv \psi \wedge \chi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then both $\sigma \mathbb{T} \psi \in \Delta$ and $\sigma \mathbb{T} \chi \in \Delta$ since the branch is complete. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma \Vdash \psi$ and $\mathfrak{M}(\Delta), \sigma \Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
If $\sigma \mathbb{F} \varphi \in \Delta$, then either $\sigma \mathbb{F} \psi \in \Delta$ or $\sigma \mathbb{F} \chi \in \Delta$ since the branch is complete. By induction hypothesis, either $\mathfrak{M}(\Delta), \sigma \not\Vdash \psi$ or $\mathfrak{M}(\Delta), \sigma \not\Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.

4. $\varphi \equiv \psi \vee \chi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then either $\sigma \mathbb{T} \psi \in \Delta$ or $\sigma \mathbb{T} \chi \in \Delta$ since the branch is complete. By induction hypothesis, either $\mathfrak{M}(\Delta), \sigma \Vdash \psi$ or $\mathfrak{M}(\Delta), \sigma \Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
- If $\sigma \mathbb{F} \varphi \in \Delta$, then both $\sigma \mathbb{F} \psi \in \Delta$ and $\sigma \mathbb{F} \chi \in \Delta$ since the branch is complete. By induction hypothesis, both $\mathfrak{M}(\Delta), \sigma \not\Vdash \psi$ and $\mathfrak{M}(\Delta), \sigma \not\Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.
5. $\varphi \equiv \psi \rightarrow \chi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then either $\sigma \mathbb{F} \psi \in \Delta$ or $\sigma \mathbb{T} \chi \in \Delta$ since the branch is complete. By induction hypothesis, either $\mathfrak{M}(\Delta), \sigma \not\Vdash \psi$ or $\mathfrak{M}(\Delta), \sigma \Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
- If $\sigma \mathbb{F} \varphi \in \Delta$, then both $\sigma \mathbb{T} \psi \in \Delta$ and $\sigma \mathbb{F} \chi \in \Delta$ since the branch is complete. By induction hypothesis, both $\mathfrak{M}(\Delta), \sigma \Vdash \psi$ and $\mathfrak{M}(\Delta), \sigma \not\Vdash \chi$. Thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.
6. $\varphi \equiv \Box \psi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then, since the branch is complete, $\sigma.n \mathbb{T} \psi \in \Delta$ for every $\sigma.n$ used on the branch, i.e., for every $\sigma' \in P(\Delta)$ such that $R\sigma\sigma'$. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma' \Vdash \psi$ for every σ' such that $R\sigma\sigma'$. Therefore, $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
- If $\sigma \mathbb{F} \varphi \in \Delta$, then for some $\sigma.n$, $\sigma.n \mathbb{F} \psi \in \Delta$ since the branch is complete. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma.n \not\Vdash \psi$. Since $R\sigma(\sigma.n)$, there is a σ' such that $\mathfrak{M}(\Delta), \sigma' \not\Vdash \psi$. Thus $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.
7. $\varphi \equiv \Diamond \psi$: If $\sigma \mathbb{T} \varphi \in \Delta$, then for some $\sigma.n$, $\sigma.n \mathbb{T} \psi \in \Delta$ since the branch is complete. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma.n \Vdash \psi$. Since $R\sigma(\sigma.n)$, there is a σ' such that $\mathfrak{M}(\Delta), \sigma' \Vdash \psi$. Thus $\mathfrak{M}(\Delta), \sigma \Vdash \varphi$.
- If $\sigma \mathbb{F} \varphi \in \Delta$, then, since the branch is complete, $\sigma.n \mathbb{F} \psi \in \Delta$ for every $\sigma.n$ used on the branch, i.e., for every $\sigma' \in P(\Delta)$ such that $R\sigma\sigma'$. By induction hypothesis, $\mathfrak{M}(\Delta), \sigma' \not\Vdash \psi$ for every σ' such that $R\sigma\sigma'$. Therefore, $\mathfrak{M}(\Delta), \sigma \not\Vdash \varphi$.

Since $\Gamma \subseteq \Delta$, $\mathfrak{M}(\Delta) \Vdash \Gamma$. □

Problem 6.4. Complete the proof of [Theorem 6.19](#).

mod:tab:cpl:
cor:entailment-completeness

Corollary 6.20. *If $\Gamma \vDash \varphi$ then $\Gamma \vdash \varphi$.*

mod:tab:cpl:
cor:weak-completeness

Corollary 6.21. *If φ is true in all models, then $\vdash \varphi$.*

6.10 Countermodels from Tableaux

mod:tab:cou:
sec

The proof of the completeness theorem doesn't just show that if $\vDash \varphi$ then $\vdash \varphi$, it also gives us a method for constructing countermodels to φ if $\not\vDash \varphi$. In the case of \mathbf{K} , this method constitutes a *decision procedure*. For suppose $\not\vDash \varphi$. Then the proof of [Proposition 6.18](#) gives a method for constructing a complete [tableau](#). The method in fact always terminates. The propositional rules for \mathbf{K} only add prefixed [formulas](#) of lower complexity, i.e., each propositional rule

[explanation](#)

need only be applied once on a branch for any signed formula $\sigma S\varphi$. New prefixes are only generated by the $\Box\mathbb{F}$ and $\Diamond\mathbb{T}$ rules, and also only have to be applied once (and produce a single new prefix). $\Box\mathbb{T}$ and $\Diamond\mathbb{F}$ have to be applied potentially multiple times, but only once per prefix, and only finitely many new prefixes are generated. So the construction either results in a closed branch or a complete branch after finitely many stages.

Once a tableau with an open complete branch is constructed, the proof of [Theorem 6.19](#) gives us an explicit model that satisfies the original set of prefixed formulas. So not only is it the case that if $\Gamma \vDash \varphi$, then a closed tableau exists and $\Gamma \vdash \varphi$, if we look for the closed tableau in the right way and end up with a “complete” tableau, we’ll not only know that $\Gamma \not\vdash \varphi$ but actually be able to construct a countermodel.

Example 6.22. We know that $\not\vdash \Box(p \vee q) \rightarrow (\Box p \vee \Box q)$. The construction of a tableau begins with:

1.	$1\mathbb{F} \Box(p \vee q) \rightarrow (\Box p \vee \Box q) \checkmark$	Assumption
2.	$1\mathbb{T} \Box(p \vee q)$	$\rightarrow\mathbb{F} 1$
3.	$1\mathbb{F} \Box p \vee \Box q \checkmark$	$\rightarrow\mathbb{F} 1$
4.	$1\mathbb{F} \Box p \checkmark$	$\vee\mathbb{F} 3$
5.	$1\mathbb{F} \Box q \checkmark$	$\vee\mathbb{F} 3$
6.	$1.1\mathbb{F} p \checkmark$	$\Box\mathbb{F} 4$
7.	$1.2\mathbb{F} q \checkmark$	$\Box\mathbb{F} 5$

The tableau is of course not finished yet. In the next step, we consider the only line without a checkmark: the prefixed formula $1\mathbb{T}\Box(p \vee q)$ on line 2. The construction of the closed tableau says to apply the $\Box\mathbb{T}$ rule for every prefix used on the branch, i.e., for both 1.1 and 1.2:

1.	$1\mathbb{F} \Box(p \vee q) \rightarrow (\Box p \vee \Box q) \checkmark$	Assumption
2.	$1\mathbb{T} \Box(p \vee q)$	$\rightarrow\mathbb{F} 1$
3.	$1\mathbb{F} \Box p \vee \Box q \checkmark$	$\rightarrow\mathbb{F} 1$
4.	$1\mathbb{F} \Box p \checkmark$	$\vee\mathbb{F} 3$
5.	$1\mathbb{F} \Box q \checkmark$	$\vee\mathbb{F} 3$
6.	$1.1\mathbb{F} p \checkmark$	$\Box\mathbb{F} 4$
7.	$1.2\mathbb{F} q \checkmark$	$\Box\mathbb{F} 5$
8.	$1.1\mathbb{T} p \vee q$	$\Box\mathbb{T} 2$
9.	$1.2\mathbb{T} p \vee q$	$\Box\mathbb{T} 2$

Now lines 2, 8, and 9, don’t have checkmarks. But no new prefix has been added, so we apply $\vee\mathbb{T}$ to lines 8 and 9, on all resulting branches (as long as they don’t close):

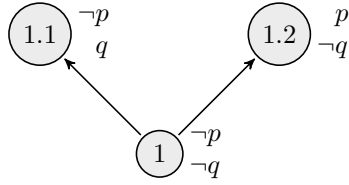


Figure 6.1: A countermodel to $\Box(p \vee q) \rightarrow (\Box p \vee \Box q)$.

mod:tab:cou:
fig:counter-Box

1.	1 F $\Box(p \vee q) \rightarrow (\Box p \vee \Box q) \checkmark$	Assumption				
2.	1 T $\Box(p \vee q)$	\rightarrow F 1				
3.	1 F $\Box p \vee \Box q \checkmark$	\rightarrow F 1				
4.	1 F $\Box p \checkmark$	\vee F 3				
5.	1 F $\Box q \checkmark$	\vee F 3				
6.	1.1 F $p \checkmark$	\Box F 4				
7.	1.2 F $q \checkmark$	\Box F 5				
8.	1.1 T $p \vee q \checkmark$	\Box T 2				
9.	1.2 T $p \vee q \checkmark$	\Box T 2				
\swarrow						
10.	<table style="display: inline-table; border: none; vertical-align: middle;"> <tr> <td style="text-align: right; padding-right: 10px;">1.1 T $p \checkmark$</td> <td style="text-align: right; padding-right: 10px;">1.1 T $q \checkmark$</td> </tr> <tr> <td style="text-align: center;">⊗</td> <td></td> </tr> </table>	1.1 T $p \checkmark$	1.1 T $q \checkmark$	⊗		\vee T 8
1.1 T $p \checkmark$	1.1 T $q \checkmark$					
⊗						
\searrow						
11.	<table style="display: inline-table; border: none; vertical-align: middle;"> <tr> <td style="text-align: right; padding-right: 10px;">1.2 T $p \checkmark$</td> <td style="text-align: right; padding-right: 10px;">1.2 T $q \checkmark$</td> </tr> <tr> <td></td> <td style="text-align: center;">⊗</td> </tr> </table>	1.2 T $p \checkmark$	1.2 T $q \checkmark$		⊗	\vee T 9
1.2 T $p \checkmark$	1.2 T $q \checkmark$					
	⊗					

There is one remaining open branch, and it is complete. From it we define the model with worlds $W = \{1, 1.1, 1.2\}$ (the only prefixes appearing on the open branch), the accessibility relation $R = \{\langle 1, 1.1 \rangle, \langle 1, 1.2 \rangle\}$, and the assignment $V(p) = \{1.2\}$ (because line 11 contains $1.2 \text{ T } p$) and $V(q) = \{1.1\}$ (because line 10 contains $1.1 \text{ T } q$). The model is pictured in Figure 6.1, and you can verify that it is a countermodel to $\Box(p \vee q) \rightarrow (\Box p \vee \Box q)$.

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Bibliography