

Modal Logic

wugouzi

March 21, 2021

Contents

1	Basic Concepts	1
1.1	Modal Languages	1
1.2	Models and Frames	3
1.3	General Frames	4
1.4	Normal Modal Logics	5
1.5	Modal Consequence Relations	6
2	Models	6
2.1	Invariance Results	6
2.1.1	Disjoint Unions	6
2.1.2	Generated submodels	6
2.1.3	Morphism for modalities	7
2.2	Bisimulations	8
2.3	Finite Models	11
2.3.1	Selecting a finite submodel	12
2.3.2	Finite models via filtrations	14
2.4	The Standard Translation	16
2.5	Modal Saturation via Ultrafilter Extensions	18
2.5.1	M-saturation	18
2.5.2	Ultrafilter extensions	19
2.6	Characterization and Definability	27
2.6.1	The van Benthem Characterization Theorem	27
2.6.2	Ultraproducts	30
2.6.3	Definability	32

3	Frames	34
3.1	Frame Definability	34
3.2	Frame Definability and Second-Order Logic	38
3.3	Definable and Undefinable Properties	43
3.4	Finite Frames	47
3.4.1	Finite transitive frames	47
3.4.2	The finite frame property	48
4	Completeness	51
4.1	Preliminaries	51
4.2	Canonical Models	54
4.3	Applications	56
4.4	Limitative Results	60
4.5	Transforming the Canonical Model	62
4.6	Step by Step	64
4.7	Rules for the Undefinable	67
4.8	Finitary Methods I	71

1 Basic Concepts

1.1 Modal Languages

Definition 1.1. The **basic modal language** is defined using a set of **proposition letters** Φ whose elements are usually denoted p, q, r and so on, and a unary modal operator \Diamond . The well-formed **formulas** ϕ of the basic modal language are given by the rule

$$\phi := p \mid \perp \mid \neg\phi \mid \psi \vee \phi \mid \Diamond\phi$$

$$\mathfrak{M}, w \Vdash \phi$$

Definition 1.2. A **modal similarity type** is a pair $\tau = (O, \rho)$ where O is a non-empty set, and ρ is a function $O \rightarrow \mathbb{N}$. The elements of O are called **modal operators**; we use $\Delta, \Delta_0, \Delta_1, \dots$ to denote elements of O . The function ρ assigns to each operator $\delta \in O$ a finite **arity**

Definition 1.3. A **modal language** $ML(\tau, \Phi)$ is built up using a modal similarity type $\tau = (O, \rho)$ and a set of proposition letters Φ . The set $Form(\tau, \Phi)$ of **modal formulas** over τ and Φ is given by the rule

$$\phi := p \mid \perp \mid \neg\phi \mid \phi_1 \vee \phi_2 \mid \Delta(\phi_1, \dots, \phi_{\rho(\Delta)})$$

where p ranges over elements of Φ

Definition 1.4. For each $\triangle \in O$ the **dual** ∇ of \triangle is defined as $\nabla(\phi_1, \dots, \phi_n) := \neg\triangle(\neg\phi_1, \dots, \neg\phi_n)$

Example 1.1 (The Basic Temporal Language). The basic temporal language is built using a set of unary operators $O = \{\langle F \rangle, \langle P \rangle\}$. The intended interpretation of a formula $\langle F \rangle\phi$ is ‘ ϕ will be true at some Future time’ and the intended interpretation of $\langle P \rangle\phi$ is ‘ ϕ was true at some Past time.’ This language is called the **basic temporal language**. Their duals are written as G and H (‘it is Going to be the case’ and ‘it always Has been the case’)

Let’s denote the converse of a relation R by R^\sim . We will call a frame of the form (T, R, R^\sim) a **bidirectional frame**, and a model built over such a frame a **bidirectional model**. From now on, we will only interpret basic temporal language in bidirectional models. That is, if $\mathfrak{M} = (T, R, R^\sim, V)$ is a bidirectional model then

$$\begin{aligned}\mathfrak{M}, t \Vdash F\phi & \quad \text{iff} \quad \exists s(Rts \wedge \mathfrak{M}, s \Vdash \phi) \\ \mathfrak{M}, t \Vdash P\phi & \quad \text{iff} \quad \exists s(R^\sim ts \wedge \mathfrak{M}, s \Vdash \phi)\end{aligned}$$

Example 1.2 (Propositional Dynamic Logic). Each of these diamonds has the form $\langle \pi \rangle$, where π denotes a (non-deterministic) **program**. The intended interpretation of $\langle \pi \rangle$ is ‘some terminating execution of π from the present state leads to a state bearing the information ϕ ’. The dual assertion $[\pi]\phi$ states that ‘every execution of π from the present state leads to a state bearing the information ϕ ’

Suppose we have fixed some set of basic programs a, b, c and so on. Then we are allowed to define complex programs π over this base as follows

1. **Choice.** If π_1, π_2 are programs, then so is $\pi_1 \cup \pi_2$
2. **Composition.** so is $\pi_1; \pi_2$
3. **Iteration.** so is π^*

This is called **regular PDL**. There are also

1. **Intersection** if π_1, π_2 are programs, then so is $\pi_1 \cap \pi_2$
execute both π_1, π_2 , in parallel
2. **test** if ϕ is a formula, then so is $\phi?$
test whether ϕ holds, and if so, continues; if not, it fails

The language of PDL has an infinite collection of diamonds, each indexed by a program π build from basic programs using the constructor $\cup, ;$ and $*$.

$$\begin{aligned} R_{\pi_1 \cup \pi_2} &= R_{\pi_1} \cup R_{\pi_2} \\ R_{\pi_1 ; \pi_2} &= R_{\pi_1} \circ R_{\pi_2} \\ R_{\pi_1^*} &= (R_{\pi_1})^* \end{aligned}$$

Suppose we have fixed a set of basic programs. Let Π be the set of programs containing the basic programs and all the programs constructed over them using the regular constructors $\cup, ;, *$. Then a **regular frame** for Π is a labeled transitive system $(W, \{R_\pi \mid \pi \in \Pi\})$ s.t. R_a is an arbitrary binary relation for each basic program a , and for all complex programs π , R_π is the binary relation inductively constructed in accordance with the previous clauses. A **regular model** for Π is a model built over a regular frame.

Example 1.3 (An Arrow Language). The type τ_{\rightarrow} of **arrow logic** is a similarity type with modal operators other than diamonds. The language of arrow logic is designed to talk about the objects in arrow structures. The well-formed formulas ϕ are given by

$$\phi := p \mid \perp \mid \neg\phi \mid \phi \vee \psi \mid \phi \circ \psi \mid \otimes\phi \mid 1'$$

$1'$ ('identity') is a nullary modality, the 'converse' operator \otimes is a diamond, and the 'composition' operator \circ is a dyadic operator. Possible readings of these operators are:

$1'$	identity	'skip'
$\otimes\phi$	converse	' ϕ conversely'
$\phi \circ \psi$	composition	'first ϕ , then ψ '

1.2 Models and Frames

Definition 1.5. A **frame** for the basic modal language is a pair $\mathfrak{F} = (W, R)$ s.t.

1. W is a non-empty set
2. R is a binary relation on W

A **model** for the basic modal language is a pair $\mathfrak{M} = (\mathfrak{F}, V)$, where \mathfrak{F} is a frame for the basic modal language and V is a function assigning to each proposition letter p in Φ a subset $V(p)$ of W . The function V is called a **valuation**. \mathfrak{M} is **based on** the frame \mathfrak{F}

Definition 1.6. Suppose w is a state in a model $\mathfrak{M} = (W, R, V)$. Then ϕ is **satisfied** in \mathfrak{M} at state w if

$$\begin{aligned} \mathfrak{M}, w \Vdash p & \text{ iff } w \in V(p), \text{ where } p \in \Phi \\ \mathfrak{M}, w \Vdash \perp & \text{ iff } \text{never} \\ \mathfrak{M}, w \Vdash \neg\phi & \text{ iff } \text{not } \mathfrak{M}, w \Vdash \phi \\ \mathfrak{M}, w \Vdash \phi \vee \psi & \text{ iff } \mathfrak{M}, w \Vdash \phi \text{ or } \mathfrak{M}, w \Vdash \psi \\ \mathfrak{M}, w \Vdash \Diamond\phi & \text{ iff } \text{for some } v \in W \text{ with } R w v \text{ we have } \mathfrak{M}, v \Vdash \phi \end{aligned}$$

It follows that $\mathfrak{M}, w \Vdash \Box\phi$ iff for all $v \in W$ s.t. $R w v$, we have $\mathfrak{M}, v \Vdash \phi$

Definition 1.7. Let τ be a modal similarity type. A τ -**frame** is a tuple \mathfrak{F} consisting of the following ingredients

1. a non-empty set W
2. for each $n \geq 0$, and each n -ary modal operator Δ in the similarity type τ , an $(n+1)$ -ary relation R_Δ

ϕ is **satisfied at a state** w in a model $\mathfrak{M} = (W, \{R_\Delta \mid \Delta \in \tau\}, V)$ when $\rho(\Delta) > 0$ if

$$\mathfrak{M}, w \Vdash \Delta(\phi_1, \dots, \phi_n) \text{ iff } \text{for some } v_1, \dots, v_n \in W \text{ with } R_\Delta w v_1 \dots v_n \text{ we have, for each } i, \mathfrak{M}, v_i \Vdash \phi_i$$

When $\rho(\Delta) = 0$ we define

$$\mathfrak{M}, w \Vdash \Delta \text{ iff } w \in R_\Delta$$

Definition 1.8. The set of all formulas that are valid in a class of frames \mathbf{F} is called the **logic** of \mathbf{F} (notation: $\Lambda_{\mathbf{F}}$)

1.3 General Frames

Definition 1.9. Given an $(n+1)$ -ary relation R on a set W , we define the following n -ary operation m_R on the power set $\mathcal{P}(W)$ of W :

$$m_R(X_1, \dots, X_n) = \{w \in W \mid R w w_1 \dots w_n \text{ for some } w_1 \in X_1, \dots, w_n \in X_n\}$$

Definition 1.10 (General Frames). Let τ be a modal similarity type. A **general τ -frame** is a pair (\mathfrak{F}, A) where $\mathfrak{F} = (W, R_\Delta)_{\Delta \in \tau}$ is a τ -frame, and A is a non-empty collection of **admissible** subsets of W closed under following operations

1. union: if $X, Y \in A$, then $X \cup Y \in A$
2. relative complement: if $X \in A$, then $W \setminus X \in A$
3. modal operations: if $X_1, \dots, X_n \in A$, then $m_{R_\Delta}(X_1, \dots, X_n) \in A$ for all $\Delta \in \tau$

A **model based on a general frame** is a triple (\mathfrak{F}, A, V) where (\mathfrak{F}, A) is a general frame and V is a valuation satisfying the constraint that for each proposition letter p , $V(p)$ is an element of A . Valuation satisfying this constraint are called **admissible** for (\mathfrak{F}, A)

A set of admissible valuations A is a 'logically closed' collection of information assignment

Definition 1.11. A formula ϕ is **valid at a state w in a general frame (\mathfrak{F}, A)** (notation: $(\mathfrak{F}, A), w \Vdash \phi$) if ϕ is true at w in every admissible model (\mathfrak{F}, A, V) on (\mathfrak{F}, A) ; and ϕ is **valid in a general frame (\mathfrak{F}, A)** (notation: $(\mathfrak{F}, A) \Vdash \phi$) if ϕ is true at every state in every admissible model

1.4 Normal Modal Logics

Definition 1.12. A **K-proof** is a finite sequence of formulas, each of which is an **axiom**, or follows from one or more earlier items in the sequence by applying a **rule of proof**. The axioms of **K** are **all instances of propositional tautologies** plus

$$(K) \quad \Box(p \rightarrow q) \rightarrow (\Box p \rightarrow \Box q)$$

$$(Dual) \quad \Diamond p \leftrightarrow \neg \Box \neg p$$

The rules of proof of **K** are

- **Modus ponens:** given ϕ and $\phi \rightarrow \psi$ prove ψ
- **Uniform substitution:** given ϕ , prove θ , where θ is obtained from ϕ by uniformly replacing proposition letters in ϕ by arbitrary formulas
- **Generalization:** given ϕ , prove $\Box \phi$

A formula ϕ is **K-provable** if it occurs as the last items of some **K-proof**, and if this is the case we write $\vdash_K \phi$

Definition 1.13 (Normal Modal Logics). A **normal modal logic Λ** is a set of formulas that contains all tautologies, $\Box(p \rightarrow q) \rightarrow (\Box p \rightarrow \Box q)$ and $\Diamond p \leftrightarrow \neg \Box \neg p$, and that is closed under modus ponens, uniform substitution and generalization. We call the smallest normal modal logic **K**

1.5 Modal Consequence Relations

Definition 1.14 (Local Semantic Consequence). Let τ be a similarity type, and let S be a class of structures of type τ (that is a class of models, a class of frames, or a class of general frames of this type). Let Σ and ϕ be a set of formulas and a single formula from a language of type τ . We say that ϕ is a **local semantic consequence of Σ over S** (notation: $\Sigma \Vdash_S \phi$) if for all models \mathfrak{M} from S , and all points w in \mathfrak{M} , if $\mathfrak{M}, w \Vdash \Sigma$ then $\mathfrak{M}, w \Vdash \phi$

Definition 1.15 (Global Semantic Consequence). Let τ, S, Σ, ϕ . ϕ is a **global semantic consequence of Σ over S** (notation $\Sigma \Vdash_S^g \phi$) iff all structures $\mathfrak{G} \in S$, if $\mathfrak{G} \Vdash \Sigma$ then $\mathfrak{G} \Vdash \phi$

2 Models

2.1 Invariance Results

Definition 2.1. Let \mathfrak{M} and \mathfrak{M}' be models of the same modal similarity type τ , and let w and w' be states in \mathfrak{M} and \mathfrak{M}' respectively. The **τ -theory** (or **τ -type**) of w is the set of all τ -formulas satisfied at w : that is, $\{\phi \mid \mathfrak{M}, w \Vdash \phi\}$. We say that w and w' are **(modally) equivalent** ($w \leftrightarrow w'$) if they have the same τ -theories

The **τ -theory** of the model \mathfrak{M} is the set of all τ -formulas satisfied by all states in fM ; that is, $\{\phi \mid \mathfrak{M} \Vdash \phi\}$. Models \mathfrak{M} and \mathfrak{M}' are called **(modally) equivalent** ($\mathfrak{M} \leftrightarrow \mathfrak{M}'$) if their theories are identical

2.1.1 Disjoint Unions

2.1.2 Generated submodels

Definition 2.2. Let $\mathfrak{M} = (W, R, V)$ and $\mathfrak{M}' = (W', R', V')$ be two models; we say that \mathfrak{M}' is a **submodel** of \mathfrak{M} if $W' \subseteq W$, R' is the restriction of R to W' , and V' is the restriction of V to \mathfrak{M}' . We say that \mathfrak{M}' is a **generated submodel** of \mathfrak{M} ($\mathfrak{M}' \twoheadrightarrow \mathfrak{M}$) if \mathfrak{M}' is a submodel of \mathfrak{M} and for all points w the following closure condition holds

$$\text{if } w \text{ is in } \mathfrak{M}' \text{ and } R w v, \text{ then } v \text{ is in } \mathfrak{M}'$$

Let \mathfrak{M} be a model, and X a subset of the domain of \mathfrak{M} ; the **submodel generated by X** is the smallest generated submodel of \mathfrak{M} whose domain contains X . A **rooted** or **point generated** model is a model that is generated by a singleton set, the element of which is called the **root** of the frame

2.1.3 Morphism for modalities

Definition 2.3 (Homomorphisms). Let τ be a modal similarity type and let \mathfrak{M} and \mathfrak{M}' be τ -models. By a **homomorphism** $f : \mathfrak{M} \rightarrow \mathfrak{M}'$, we mean a function $f : W \rightarrow W'$ satisfying

1. For each proposition letter p and each element w from \mathfrak{M} , if $w \in V(p)$, then $f(w) \in V'(p)$
2. For each $n \geq 0$ and each n -ary $\triangle \in \tau$ and $(n+1)$ -tuple \bar{w} from \mathfrak{M} , if $(w_0, \dots, w_n) \in R_\triangle$, then $(f(w_0), \dots, f(w_n)) \in R'_\triangle$ (the **homomorphic condition**)

Definition 2.4 (Strong Homomorphisms, Embeddings and Isomorphisms). Let τ be a modal similarity type and let \mathfrak{M} and \mathfrak{M}' be τ -models. By a **strong homomorphism** $f : \mathfrak{M} \rightarrow \mathfrak{M}'$, we mean a function $f : W \rightarrow W'$ satisfying

1. For each proposition letter p and each element w from \mathfrak{M} iff $w \in V(p)$, then $f(w) \in V'(p)$
2. For each $n \geq 0$ and each n -ary $\triangle \in \tau$ and $(n+1)$ -tuple \bar{w} from \mathfrak{M} iff $(w_0, \dots, w_n) \in R_\triangle$, iff $(f(w_0), \dots, f(w_n)) \in R'_\triangle$ (the **strong homomorphic condition**)

An **embedding** of \mathfrak{M} into \mathfrak{M}' is a strong homomorphism $f : \mathfrak{M} \rightarrow \mathfrak{M}'$ which is injective. An **isomorphism** is a bijective strong homomorphism

Proposition 2.5. *Let τ be a modal similarity type and let \mathfrak{M} and \mathfrak{M}' be τ -models. Then the following holds*

1. *for all elements w and w' of \mathfrak{M} and \mathfrak{M}' , respectively, if there exists a surjective strong homomorphism $f : \mathfrak{M} \rightarrow \mathfrak{M}'$ with $f(w) = w'$, then w and w' are modally equivalent*
2. *If $\mathfrak{M} \cong \mathfrak{M}'$, then $\mathfrak{M} \leftrightarrow \mathfrak{M}'$*

Definition 2.6 (Bounded Morphisms - the Basic Case). Let \mathfrak{M} and \mathfrak{M}' be models for the basic modal language. A mapping $f : \mathfrak{M} = (W, R, V) \rightarrow \mathfrak{M}' = (W', R', V')$ is a **bounded morphism** if it satisfies

1. w and $f(w)$ satisfy the same proposition letters
2. f is a homomorphism w.r.t. the relation R (if Rwv then $R'f(w)f(v)$)
3. If $R'f(w)v'$ then there exists v s.t. Rwv and $f(v) = v'$ (the **back condition**)

If there is a **surjective** bounded morphism from \mathfrak{M} to \mathfrak{M}' , then we say that \mathfrak{M}' is a **bounded morphic image** of \mathfrak{M} , and write $\mathfrak{M} \twoheadrightarrow \mathfrak{M}'$

Proposition 2.7. *Let τ be a modal similarity type and let \mathfrak{M} and \mathfrak{M}' be τ -models s.t. $f : \mathfrak{M} \rightarrow \mathfrak{M}'$ is a bounded morphism. Then for each modal formula ϕ , and each element w of \mathfrak{M} we have $\mathfrak{M}, w \models \phi$ iff $\mathfrak{M}', f(w) \models \phi$.*

Let τ be a modal similarity type containing only diamonds (thus if \mathfrak{M} is a τ -model, it has the form (W, R_1, \dots, V) where each R_i is a binary relation on W). In this context we will call a τ -model \mathfrak{M} **tree-like** if the structure $(W, \bigcup_i R_i, V)$ is a tree

Proposition 2.8. *Assume that τ is a modal similarity type containing only diamonds. Then for any rooted τ -models \mathfrak{M} there exists a tree-like τ -models \mathfrak{M}' s.t. $\mathfrak{M}' \twoheadrightarrow \mathfrak{M}$. Hence any satisfiable τ -formula is satisfiable in a tree-like model*

Proof. Let w be the root of \mathfrak{M} . Define the model \mathfrak{M}' as follows. Its domain W' consist of all finite sequences (w, u_1, \dots, u_n) s.t. $n \geq 0$ and for some modal operators $\langle a_1 \rangle, \dots, \langle a_n \rangle \in \tau$ there is a path $w R_{a_1} u_1 \dots R_{a_n} u_n$ in \mathfrak{M} . Define $(w, u_1, \dots, u_n) R'_a (w, v_1, \dots, w_m)$ to hold if $m = n + 1, u_i = v_i$ for $i = 1, \dots, n$ and $R_a u_n v_m$ holds in \mathfrak{M} . That is, R'_a relates two sequences iff the second is an extension of the first with a state from \mathfrak{M} that is a successor of the last element of the first sequence. Finally, V' is defined by putting $(w, u_1, \dots, u_n) \in V'(p)$ iff $u_n \in V(p)$. The mapping $f : (w, u_1, \dots, u_n) \mapsto u_n$ defines a surjective bounded morphism from \mathfrak{M}' to \mathfrak{M} \square

2.2 Bisimulations

Definition 2.9 (Bisimulation - the Basic Case). Let $\mathfrak{M} = (W, R, V)$ and $\mathfrak{M}' = (W', R', V')$ be two models

A non-empty binary relation $Z \subseteq W \times W'$ is called a **bisimulation between** \mathfrak{M} and \mathfrak{M}' (notation: $Z : \mathfrak{M} \rightleftharpoons \mathfrak{M}'$) if

1. If wZw' then w and w' satisfy the same proposition letters
2. If wZw' and Rwv , then there exists v' (in \mathfrak{M}') s.t. vZv' and $R'w'v'$ (the **forth condition**)
3. The converse of (2): if wZw' and $R'w'v'$, then there exists v (in \mathfrak{M}) s.t. vZv' and Rwv (the **back condition**)

When Z is a bisimulation linking two states w in \mathfrak{M} and w' in \mathfrak{M}' we say that w and w' are **bisimilar**, and we write $Z : \mathfrak{M}, w \rightleftharpoons \mathfrak{M}', w'$. If there is a bisimulation, we sometimes write $\mathfrak{M}, w \rightleftharpoons \mathfrak{M}', w'$ or $w \rightleftharpoons w'$

Definition 2.10 (Bisimulation - the General Case). Let τ be a modal similarity type, and let $\mathfrak{M} = (W, R_\Delta, V)_{\Delta \in \tau}$ and $\mathfrak{M}' = (W', R'_\Delta, V')_{\Delta \in \tau}$ be τ -models. A non-empty binary relation $Z \subseteq W \times W'$ is called a **bisimulation** between \mathfrak{M} and \mathfrak{M}' ($Z : \mathfrak{M} \rightleftharpoons \mathfrak{M}'$) if the above condition 1 is satisfied and

2. If wZw' and $R_\Delta wv_1 \dots v_n$ then there are $v'_1, \dots, v'_n \in W'$ s.t. $R'_\Delta w'v'_1 \dots v'_n$ and for all i ($1 \leq i \leq n$) $v_iZv'_i$ (the **forth** condition)
3. If wZw' and $R'_\Delta w'v'_1 \dots v'_n$ then there are $v_1, \dots, v_n \in W$ s.t. $R_\Delta wv_1 \dots v_n$ and for all i ($1 \leq i \leq n$) $v_iZv'_i$ (the **back** condition)

Proposition 2.11. Let τ be a modal similarity type, and let $\mathfrak{M}, \mathfrak{M}'$ and \mathfrak{M}_i ($i \in I$) be τ -models

1. If $\mathfrak{M} \cong \mathfrak{M}'$, then $\mathfrak{M} \rightleftharpoons \mathfrak{M}'$
2. For every $i \in I$, and every w in \mathfrak{M}_i , $\mathfrak{M}_i, w \rightleftharpoons \biguplus_i \mathfrak{M}_i, w$
3. If $\mathfrak{M}' \succcurlyeq \mathfrak{M}$, then $\mathfrak{M}', w \rightleftharpoons \mathfrak{M}, w$ for all w in \mathfrak{M}'
4. If $f : \mathfrak{M} \rightarrow \mathfrak{M}'$, then $\mathfrak{M}, w \rightleftharpoons \mathfrak{M}', f(w)$ for all w in \mathfrak{M}

Proof. Suppose $\mathfrak{M} = (W, R_\Delta, V)_{\Delta \in \tau}$ and $\mathfrak{M}' = (W', R'_\Delta, V')_{\Delta \in \tau}$ $\mathfrak{M}_i \subseteq \biguplus_i \mathfrak{M}_i$

1. Suppose $f : \mathfrak{M} \cong \mathfrak{M}'$, then we define wZw' iff $w' = f(w)$ where $w \in W, w' \in W'$. Bisimulation comes from the definition of the isomorphism
2. Define the relation $Z = \{(w, w) \mid w \in \mathfrak{M}_i\} \subseteq \mathfrak{M}_i \times \biguplus_i \mathfrak{M}_i$. The first condition comes from the invariance. The forth condition is obvious. For the back condition, if $R'_\Delta w'v'_1 \dots v'_n$ and $w' \in W$, then $v'_1, \dots, v'_n \in W$ since each $R_{\Delta, i}$ is disjoint and we have $R_{\Delta, i}w'v'_1 \dots v'_n$
3. Define the relation $Z = \{(w, w) \mid w \in \mathfrak{M}'\} \subseteq \mathfrak{M}' \times \mathfrak{M}$. The first condition comes from the invariance. Forth condition is obvious. For the back condition, suppose wZw and $R'_\Delta wv'_1 \dots v'_n$, by the definition, $v'_1, \dots, v'_n \in W$ and $R_\Delta wv'_1 \dots v'_n$
4. Define $Z = \{(w, f(w)) \mid w \in W\}$. The first condition comes from the definition. If wZw' and $R_\Delta wv_1 \dots v_n$, then $R'_\Delta f(w)f(v_1) \dots f(v_n)$. If wZw' and $R'_\Delta w'v'_1 \dots v'_n$, then there is v_1, \dots, v_n s.t. $R_\Delta wv_1, \dots, v_n$ and $f(v_i) = v'_i$ for $1 \leq i \leq n$

□

Theorem 2.12. *Let τ be a modal similarity type, and let $\mathfrak{M}, \mathfrak{M}'$ be τ -models. Then, for every $w \in W$ and $w' \in W'$, $w \simeq w'$ implies that $w \leftrightarrow w'$. In other words, modal formulas are invariant under bisimulation*

Proof. Induction on the complexity of ϕ .

Suppose ϕ is $\Diamond\psi$, we have $\mathfrak{M}, w \Vdash \Diamond\psi$ iff there exists a v in \mathfrak{M} s.t. Rwv and $\mathfrak{M}, v \Vdash \psi$. As $w \simeq w'$, there exists a v' in \mathfrak{M}' s.t. $R'w'v'$ and $v \simeq v'$. By the I.H., $\mathfrak{M}', v' \Vdash \psi$, hence $\mathfrak{M}', w' \Vdash \Diamond\psi$ □

Example 2.1 (Bisimulation and First-Order Logic).

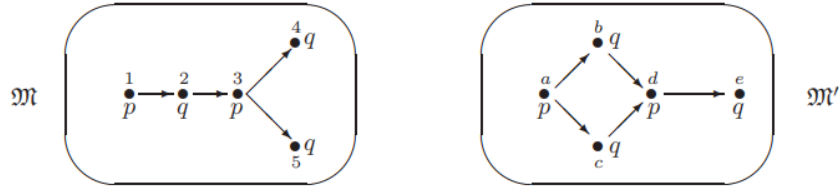


Fig. 2.4. Bisimilar models.

Example 2.2.



Fig. 2.5. Equivalent but not bisimilar.

\mathfrak{M} is **image-finite** if for each state u in \mathfrak{M} and each relation R in \mathfrak{M} , the set $\{(v_1, \dots, v_n) \mid Ruv_1 \dots v_n\}$ is finite

Theorem 2.13 (Hennessy-Milner Theorem). *Let τ be a modal similarity type and let \mathfrak{M} and \mathfrak{M}' be two image-finite τ -models. Then for every $w \in W$ and $w' \in W'$, $w \simeq w'$ iff $w \leftrightarrow w'$*

Proof. Assume that our similarity type τ only contains a single diamond. The direction from left to right follows from Theorem 2.12

Suppose $w \leftrightarrow w'$. The first condition is immediate. If Rwv , assume there is no v' in \mathfrak{M}' with $R'w'v'$ and $v \leftrightarrow v'$. Let $S' = \{u' \mid R'w'u'\}$. Note that S' must be non-empty, for otherwise $\mathfrak{M}', w' \models \Box\perp$, which would contradict $w \leftrightarrow w'$ since $\mathfrak{M}, w \models \Diamond\top$. Furthermore, as \mathfrak{M}' is image-finite, S' must be finite, say $S' = \{w'_1, \dots, w'_n\}$. By assumption, for every $w'_i \in S'$ there exists a formula ψ_i s.t. $\mathfrak{M}, v \models \psi_i$, but $\mathfrak{M}', w'_i \not\models \psi_i$. It follows that

$$\mathfrak{M}, w \models \Diamond(\psi_1 \wedge \dots \wedge \psi_n) \quad \text{and} \quad \mathfrak{M}', w' \not\models \Diamond(\psi_1 \wedge \dots \wedge \psi_n)$$

□

Exercise 2.2.1. Suppose that $\{Z_i \mid i \in I\}$ is a non-empty collection of bisimulations between \mathfrak{M} and \mathfrak{M}' . Prove that the relation $\bigcup_{i \in I} Z_i$ is also a bisimulation between \mathfrak{M} and \mathfrak{M}' . Conclude that if \mathfrak{M} and \mathfrak{M}' are bisimilar, then there is a maximal bisimulation between \mathfrak{M} and \mathfrak{M}' .

Proof. 1. If $(w, w') \in \bigcup_{i \in I} Z_i$, then $(w, w') \in Z_j$ for some $j \in I$ and hence they satisfy the same propositional letters

2. If $(w, w') \in \bigcup_{i \in I} Z_i$ and $R_\Delta wv_1 \dots v_n$, since $(w, w') \in Z_j$ for some $j \in I$, we have $R'_\Delta w'v'_1 \dots v'_n$ and $v_i Z_j v'_i$ for all $1 \leq i \leq n$, which means $(v_i, v'_i) \in \bigcup_{i \in I} Z_i$ for all $1 \leq i \leq n$

3. similarly

□

Remark (Bisimulations for the Basic Temporal Language and Arrow Logic). When working with the basic temporal language, we usually work with models (W, R, V) and implicitly take R_p to be R^\sim . Thus we need a notion of bisimulation between models (W, R, V) and (W', R', V') to be a relation Z between the states of the two models that satisfies the clauses of Definition 2.9, and in addition the following

4. If wZw' and Rvw , then there exists v' in \mathfrak{M}' s.t. vZv' and $R'v'w'$

5. Converse of 4: if wZw' and $R'v'w'$, then there exists v in \mathfrak{M} s.t. vZv'

2.3 Finite Models

Definition 2.14 (Finite Model Property). Let τ be a modal similarity type, and let M be a class of τ -models. We say that τ has the **finite model property w.r.t.** M if the following holds: if ϕ is a formula of similarity type τ , and ϕ is satisfiable in some model in M , then ϕ is satisfiable in a **finite** model in M

2.3.1 Selecting a finite submodel

Definition 2.15 (Degree). We define the **degree** of modal formulas as follows:

$$\begin{aligned}
\deg(p) &= 0 \\
\deg(\perp) &= 0 \\
\deg(\neg\phi) &= \deg(\phi) \\
\deg(\phi \vee \psi) &= \max\{\deg(\phi), \deg(\psi)\} \\
\deg(\Delta(\phi_1, \dots, \phi_n)) &= 1 + \max\{\deg(\phi_1), \dots, \deg(\phi_n)\}
\end{aligned}$$

Proposition 2.16. Let τ be a finite modal similarity type, and assume our collection of proposition letters is finite as well

1. for all n , up to logical equivalence there are only finitely many formulas of degree at most n
2. for all n , and every τ -model \mathfrak{M} and state w of \mathfrak{M} , the set of all τ -formulas of degree at most n that are satisfied by w , is equivalent to a single formula

Definition 2.17 (n -Bisimulation). Let \mathfrak{M} and \mathfrak{M}' be models, and let w and w' be states of \mathfrak{M} and \mathfrak{M}' , respectively. We say that w and w' are **n -bisimilar** ($w \rightleftharpoons_n w'$) if there exists a sequence of binary relations $Z_n \subseteq \dots \subseteq Z_0$ with the following properties (for $i + 1 \leq n$)

1. wZ_nw'
2. if vZ_0v' then v and v' agree on all proposition letters
3. if $vZ_{i+1}v'$ and Rvu then there exists u' with $R'v'u'$ and uZ_iv'
4. if $vZ_{i+1}v'$ and $R'v'u'$, then there exists u with Rvu and uZ_iv'

Proposition 2.18. Let τ be a finite modal similarity type, Φ a finite set of proposition letters, and let \mathfrak{M} and \mathfrak{M}' be models for this language. Then for every w in \mathfrak{M} and w' in \mathfrak{M}' , the following are equivalent

1. $w \rightleftharpoons_n w'$
2. w and w' agree on all modal formulas of degree at most n .

Proof. $2 \rightarrow 1$. if $n = 0$, obvious.

If $n = k$ and the proposition holds. Now suppose $n = k + 1$. Now w and w' agree on all modal formulas of degree at most $n + 1$. If there is not v, v'

s.t. v and v' agree on all modal formulas of degree at most n and Rwv and Rwv' . Let $S' = \{u' \mid R'u'u'\}$ and S' is finite, say $S' = \{w'_1, \dots, w'_n\}$. By assumption, for every $w'_i \in S'$ there exists a formula ψ_i of degree at most n s.t. $\mathfrak{M}, v \models \psi_i$ but $\mathfrak{M}', w'_i \not\models \psi_i$. It follows that

$$\mathfrak{M}, w \models \diamond(\psi_1 \wedge \dots \wedge \psi_n) \text{ and } \mathfrak{M}', w' \not\models \diamond(\psi_1 \wedge \dots \wedge \psi_n)$$

□

Definition 2.19. Let τ be a modal similarity type containing only diamonds. Let $\mathfrak{M} = (W, R_1, \dots, R_n, \dots, V)$ be a rooted τ -model with root w . The notion of the **height** of states in \mathfrak{M} is defined by induction.

The only element of height 0 is the root of the model; the states of height $n + 1$ are those immediate successors of elements of height n that have not yet assigned a height smaller than $n + 1$. The **height of a model** \mathfrak{M} is the maximum n s.t. there is a state of height n in \mathfrak{M} , if such a maximum exists; otherwise the height of \mathfrak{M} is infinite

For a natural number k , the **restriction** of \mathfrak{M} to k ($\mathfrak{M} \upharpoonright k$) is defined as the submodel containing only states whose height is at most k . $(\mathfrak{M} \upharpoonright k) = (W_k, R_{1k}, \dots, R_{nk}, \dots, V_k)$, where $W_k = \{v \mid \text{height}(v) \leq k\}$, $R_{nk} = R_n \cap (W_k \times W_k)$, and for each p , $V_k(p) = V(p) \cap W_k$

Lemma 2.20. Let τ be a modal similarity type that contains only diamonds. Let \mathfrak{M} be a rooted τ -model, and let k be a natural number. Then for every state w of $(\mathfrak{M} \upharpoonright k)$, we have $(\mathfrak{M} \upharpoonright k), w \simeq_l \mathfrak{M}, w$, where $l = k - \text{height}(w)$

Theorem 2.21 (Finite Model Property - via Selection). Let τ be a modal similarity type containing only diamonds, and let ϕ be a τ -formula. If ϕ is satisfiable, then it is satisfiable on a finite model

Proof. Fix a modal formula ϕ with $\deg(\phi) = k$. We restrict our modal similarity type τ and our collection of proposition letters to the modal operators and proposition letters actually occurring in ϕ . Let \mathfrak{M}_1, w_1 be s.t. $\mathfrak{M}_1, w_1 \models \phi$. By Proposition 2.8, there exists a tree-like model \mathfrak{M}_2 with root w_2 s.t. $\mathfrak{M}_2, w_2 \models \phi$. Let $\mathfrak{M}_3 := (\mathfrak{M}_2 \upharpoonright k)$. By Lemma 2.20 we have $\mathfrak{M}_2, w_2 \simeq_k \mathfrak{M}_3, w_2$ and by Proposition 2.18 it follows that $\mathfrak{M}_3, w_2 \models \phi$

By induction on $n \leq k$ we define finite sets of states S_0, \dots, S_k and a (final) model \mathfrak{M}_4 with domain $S_0 \cup \dots \cup S_k$; the points in each S_n will have height n

Define S_0 to be the singleton $\{w_2\}$. Next, assume that S_0, \dots, S_n have already been defined. Fix an element v of S_n . By Proposition 2.16 there are only finitely many non-equivalent modal formulas whose degree is at most

$k - n$, say ψ_1, \dots, ψ_m . For each formula that is of the form $\langle a \rangle \chi$ and holds in \mathfrak{M}_3 at v , select a state u from \mathfrak{M}_3 s.t. $R_a v u$ and $\mathfrak{M}_3, u \models \chi$. Add all these u s to S_{n+1} , and repeat this selection process for every state in S_n . S_{n+1} is defined as the set of all points that have been selected in this way

Finally, define \mathfrak{M}_4 as follows. Its domain is $S_0 \cup \dots \cup S_k$; as each S_i is finite, \mathfrak{M}_4 is finite. The relations and valuation are obtained by restricting the relations and valuations of \mathfrak{M}_3 to the domain of \mathfrak{M}_4 \square

2.3.2 Finite models via filtrations

Definition 2.22. A set of formulas Σ is **closed under subformulas** (or **subformula closed**) if for all formulas ϕ, ϕ' : if $\phi \vee \phi' \in \Sigma$ then so are ϕ and ϕ' ; if $\neg\phi \in \Sigma$ then so is ϕ ; and if $\triangle(\phi_1, \dots, \phi_n) \in \Sigma$ then so are ϕ_1, \dots, ϕ_n

Definition 2.23 (Filtrations). We work in the basic modal language. Let $\mathfrak{M} = (W, R, V)$ be a model and Σ a subformula closed set of formulas. Let \leftrightarrow_Σ be the relation on the states of \mathfrak{M} defined by

$$w \leftrightarrow_\Sigma v \text{ iff for all } \phi \in \Sigma : (\mathfrak{M}, w \models \phi \text{ iff } \mathfrak{M}, v \models \phi)$$

Note that \leftrightarrow_Σ is an equivalence relation. We denote the equivalence class of a state w of \mathfrak{M} w.r.t. \leftrightarrow_Σ by $|w|_\Sigma$, or simply $|w|$. The mapping $w \mapsto |w|$ is called the **natural map**

Let $W_\Sigma = \{|w|_\Sigma \mid w \in W\}$. Suppose \mathfrak{M}_Σ^f is any model (W^f, R^f, V^f) s.t.

1. $W^f = W_\Sigma$
2. if $R w v$ then $R^f |w| |v|$
3. if $R^f |w| |v|$ then for all $\diamond\phi \in \Sigma$, if $\mathfrak{M}, v \models \phi$ then $\mathfrak{M}, w \models \diamond\phi$
4. $V^f(p) = \{|w| \mid \mathfrak{M}, w \models p\}$, for all proposition letters p in Σ

\mathfrak{M}_Σ^f is called a **filtration of \mathfrak{M} through Σ** ; we will often suppress subscripts and write \mathfrak{M}^f instead of \mathfrak{M}_Σ^f

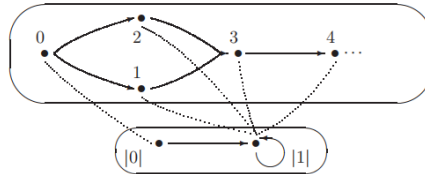


Fig. 2.6. A model and its filtration.

Let $\mathfrak{M} = (\mathbb{N}, R, V)$, where $R = \{(0, 1), (0, 2), (1, 3)\} \cup \{(n, n+1) \mid n \geq 2\}$, and V has $V(p) = \mathbb{N} \setminus \{0\}$ and $V(q) = \{2\}$

Further assume $\Sigma = \{\Diamond p, p\}$. Σ is subformula closed. Then, the model $\mathfrak{N} = (\{|0|, |1|\}, \{(|0|, |1|), (|1|, |1|)\}, V')$, where $V'(p) = \{|1|\}$ is a filtration of \mathfrak{M} through Σ . \mathfrak{N} is not a bounded morphic image of \mathfrak{M} : any bounded morphism would have to preserve the formula q

Proposition 2.24. *Let Σ be a finite subformula closed set of basic modal formulas. For any model \mathfrak{M} , if \mathfrak{M}^f is a filtration of \mathfrak{M} through a subformula closed set Σ , then \mathfrak{M}^f contains at most 2^n nodes (where n denotes the size of Σ)*

Proof. The states of \mathfrak{M}^f are the equivalence classes in W_Σ . Let g be the function with domain W_Σ and range $\mathcal{P}(\Sigma)$ defined by $g(|w|) = \{\phi \in \Sigma \mid \mathfrak{M}, w \Vdash \phi\}$. It follows from the definition of \sim_Σ that g is well defined and injective. Thus $|W_\Sigma| \leq 2^n, n = |\Sigma|$ \square

Theorem 2.25 (Filtration Theorem). *Consider the basic modal language. Let $\mathfrak{M}^f = (W_\Sigma, R^f, V^f)$ be a filtration of \mathfrak{M} through a subformula closed set Σ . Then for all formulas $\phi \in \Sigma$, and all nodes w in \mathfrak{M} , we have $\mathfrak{M}, w \Vdash \phi$ iff $\mathfrak{M}^f, |w| \Vdash \phi$*

Proof. Suppose $\Diamond \phi \in \Sigma$ and $\mathfrak{M}, w \Vdash \Diamond \phi$. Then there is a v s.t. Rwv and $\mathfrak{M}, v \Vdash \phi$. As \mathfrak{M}^f is a filtration, $R^f|w||v|$. As Σ is a subformula closed, $\phi \in \Sigma$, thus by the inductive hypothesis $\mathfrak{M}^f, |v| \Vdash \phi$. Hence $\mathfrak{M}^f, |w| \Vdash \Diamond \phi$

Suppose $\Diamond \phi \in \Sigma$ and $\mathfrak{M}^f, |w| \Vdash \Diamond \phi$. Thus there is a state $|v|$ in \mathfrak{M}^f s.t. $R^f|w||v|$ and $\mathfrak{M}^f, |v| \Vdash \phi$. As $\phi \in \Sigma$, we have $\mathfrak{M}, v \Vdash \phi$. By the definition, we have $\mathfrak{M}, w \Vdash \Diamond \phi$ \square

Note that clauses 2 and 3 of Definition 2.3.2 are designed to make the modal case of the inductive step go through.

Define

1. $R^s|w||v|$ iff $\exists w' \in |w| \exists v' \in |v| Rw'v'$
2. $R^l|w||v|$ iff for all formulas $\Diamond \phi \in \Sigma$: $\mathfrak{M}, v \Vdash \phi$ implies $\mathfrak{M}, w \Vdash \Diamond \phi$

These relations give rise to the **smallest** and **largest** filtrations respectively

Lemma 2.26. *Consider the basic modal language. Let \mathfrak{M} be any model, Σ any subformula closed set of formulas, W_Σ the set of equivalence classes induced by \sim_Σ , and V^f the standard valuation on W_Σ . Then both (W_Σ, R^s, V^f) and (W_Σ, R^l, V^f) are filtrations of \mathfrak{M} through Σ . Furthermore, if (W_Σ, R^f, V^f) is any filtration of \mathfrak{M} through Σ , then $R^s \subseteq R^f \subseteq R^l$*

Proof. If Rwv , if $\mathfrak{M}, v \Vdash \phi$, then $\mathfrak{M}, w \Vdash \Diamond\phi$, hence $R^l|w||v|$

For any (W_Σ, R^f, V^f) . $R^s \subseteq R^f$ by clause 2. $R^f \subseteq R^l$ by clause 2 \square

Theorem 2.27 (Finite Model Property - via Filtrations). *Let ϕ be a basic modal formula. if ϕ is satisfiable, then it is satisfiable on a finite model. Indeed, it is satisfiable on a finite model containing at most 2^m nodes, where m is the number of subformulas of ϕ*

Proof. Assume that ϕ is satisfiable on a model \mathfrak{M} ; take any filtration of \mathfrak{M} through the set of subformulas . \square

Lemma 2.28. *Let \mathfrak{M} be a model, Σ a subformula closed set of formulas, and W_Σ the set of equivalence classes induced on \mathfrak{M} by \leftrightarrow_Σ . Let R^t be the binary relation on W_Σ defined by*

$$R^t|w||v| \text{ iff for all } \phi, \text{ if } \Diamond\phi \in \Sigma \text{ and } \mathfrak{M}, v \Vdash \phi \vee \Diamond\phi \text{ then } \mathfrak{M}, w \Vdash \Diamond\phi$$

If R is transitive then (W_Σ, R^t, V^f) is a filtration and R^t is transitive

Definition 2.29. Let (W, R, V) be a transitive frame. A **cluster** on (W, R, V) is a maximal, nonempty equivalence class under R . That is, $C \subseteq W$ is a cluster if the restriction of R to C is an equivalence relation

A cluster is **simple** if it consists of a single reflexive point, and **proper** if it consists more than one point

2.4 The Standard Translation

Definition 2.30. For τ a modal similarity type and Φ a collection of proposition letters, let $\mathcal{L}_\tau^1(\Phi)$ be the first-order language (with equality) which has unary predicates P_0, P_1, \dots corresponding to the proposition letters p_0, p_1, \dots in Φ , and an $(n+1)$ -ary relation symbol R_Δ for each $(n\text{-ary})$ modal operator Δ in our similarity type. We write $\alpha(x)$ to denote a first-order formula α with one free variable, x

Definition 2.31 (Standard Translation). Let x be a first-order variable. The **standard translation** ST_x taking modal formulas to first-order formulas in

$\mathcal{L}_\tau^1(\Phi)$ is defined as

$$\begin{aligned}
ST_x(p) &= Px \\
ST_x(\perp) &= x \neq x \\
ST_x(\neg\phi) &= \neg ST_x(\phi) \\
ST_x(\phi \vee \psi) &= ST_x(\phi) \vee ST_x(\psi) \\
ST_x(\Delta(\phi_1, \dots, \phi_n)) &= \exists y_1 \dots \exists y_n (R_\Delta x y_1 \dots y_n \wedge \\
&\quad ST_{y_1}(\phi_1) \wedge \dots \wedge ST_{y_n}(\phi_n))
\end{aligned}$$

where y_1, \dots, y_n are fresh variables.

$$ST_x(\diamond\phi) = \exists y (Rxy \wedge ST_y(\phi))$$

$$ST_x(\Box\phi) = \forall y (Rxy \rightarrow ST_y(\phi))$$

Proposition 2.32 (Local and Global Correspondence on Models). *Fix a modal similarity type τ , and let ϕ be a τ -formula. Then*

1. *For all \mathfrak{M} and all states w of \mathfrak{M} : $\mathfrak{M}, w \Vdash \phi$ iff $\mathfrak{M} \models ST_x(\phi)[w]$*
2. *For all \mathfrak{M} : $\mathfrak{M} \Vdash \phi$ iff $\mathfrak{M} \models \forall x ST_x(\phi)$*

Proposition 2.33. 1. *Let τ be a modal similarity type that only contains diamonds. Then, every τ -formula ϕ is equivalent to a first-order formula containing at most two variables*

2. *If τ does not contain modal operators Δ whose arity exceeds n , all τ -formulas are equivalent to first-order formulas containing at most $(n + 1)$ variables*

Proof. Assume τ contains only diamonds $\langle a \rangle, \langle b \rangle$. Fix two distinct variables x and y . Define two variants ST_x and ST_y of the standard translation as follows

$$\begin{aligned}
ST_x(p) &= Px & ST_y(p) &= Py \\
ST_x(\perp) &= x \neq x & ST_y(\perp) &= y \neq y \\
ST_x(\neg\phi) &= \neg ST_x(\phi) & ST_y(\neg\phi) &= \neg ST_y(\phi) \\
ST_x(\phi \vee \psi) &= ST_x(\phi) \vee ST_x(\psi) & ST_y(\phi \vee \psi) &= ST_y(\phi) \vee ST_y(\psi) \\
ST_x(\langle a \rangle\phi) &= \exists y (R_a x y \wedge ST_y(\phi)) & ST_y(\langle a \rangle\phi) &= \exists x (R_a y x \wedge ST_x(\phi))
\end{aligned}$$

Then for any τ -formula ϕ , its ST_x -translation contains at most the two variables x and y , and $ST_x(\phi)$ is equivalent to the original standard translation of ϕ \square

Example 2.3.

$$\begin{aligned}
ST_x(\diamond(\Box p \rightarrow q)) &= \exists y(Rxy \wedge ST_y(\Box p \rightarrow q)) \\
&= \exists y(Rxy \wedge (\forall x(Ryx \rightarrow ST_x(p)) \rightarrow Qy)) \\
&= \exists y(Rxy \wedge (\forall x(Ryx \rightarrow Px) \rightarrow Qy))
\end{aligned}$$

Rxx is not equivalent to any modal formula. Suppose ϕ is a modal formula s.t. $ST_x(\phi)$ is equivalent to Rxx . Let \mathfrak{M} be a singleton reflexive model and let w be the unique state in \mathfrak{M} ; obviously $\mathfrak{M} \models Rxx[w]$. Let \mathfrak{N} be a model based on the strict ordering of the integers; for every integer v , $\mathfrak{N} \models \neg Rxx[v]$. Let Z be the relation which links every integer with the unique state in fM , and assume that the valuations in \mathfrak{N} and \mathfrak{M} are s.t. Z is a bisimulation.

$$\mathfrak{M} \models Rxx[w] \Rightarrow \mathfrak{M}, w \Vdash \phi \Rightarrow \mathfrak{N}, v \Vdash \phi \Rightarrow \mathfrak{N} \models Rxx[v]$$

Definition 2.34. Let τ be a modal similarity type, C a class of τ -models, and Γ a set of formulas over τ . We say that Γ **defines** or **characterizes** a class K of models **within** C if for all models \mathfrak{M} in C we have that \mathfrak{M} is in K iff $\mathfrak{M} \models \Gamma$. If C is the class of all τ -models, we simply say that Γ defines or characterizes K ; we omit brackets whenever Γ is a singleton. We say that a formula ϕ defines a **property** whenever ϕ defines the class of models satisfying the property

2.5 Modal Saturation via Ultrafilter Extensions

2.5.1 M-saturation

Definition 2.35 (Hennessy-Milner Classes). Let τ be a modal similarity type, and K a class of τ -models. K is a **Hennessy-Milner class**, or **has the Hennessy-Milner property**, if for every two models \mathfrak{M} and \mathfrak{M}' in K and any two states w, w' of \mathfrak{M} and \mathfrak{M}' , respectively, $w \leftrightarrow w'$ implies $\mathfrak{M}, w \rightleftharpoons \mathfrak{M}', w'$

For example, by Theorem 2.13 the class of image-finite models has the Hennessy-Milner property.

Suppose we are working in the basic modal language. Let $\mathfrak{M} = (W, R, V)$ be a model, let w be a state in W and let $\Sigma = \{\phi_0, \phi_1, \dots\}$ be an infinite set of formulas. Suppose that w has successors v_0, v_1, \dots , where respectively $\phi_0, \phi_0 \wedge \phi_1, \phi_0 \wedge \phi_1 \wedge \phi_2, \dots$ hold. If there is no successor v of w where **all** formulas from Σ hold **at the same time**, then the model is in some sense incomplete. A model is called **m-saturated** if incompleteness of this kind does not occur

Suppose that we are looking for a successor of w at which every formula ϕ_i of the infinite set of formulas $\Sigma = \{\phi_0, \phi_1, \dots\}$ holds. M-saturation is a

kind of compactness property, according to which it suffices to find satisfying successors of w for arbitrary finite approximations of Σ

Definition 2.36 (M-saturation). Let $\mathfrak{M} = (W, R, V)$ be a model of the basic modal similarity type, X a subset of W and Σ a set of modal formulas. Σ is **satisfiable** in the set X if there is a state $x \in X$ s.t. $\mathfrak{M}, x \Vdash \phi$ for all $\phi \in \Sigma$. Σ is **finitely satisfiable** in X if every finite subset of Σ is satisfiable in X

The model \mathfrak{M} is called **m -saturated** if it satisfies the following condition for every state $w \in W$ and every set Σ of modal formulas:

If Σ is finitely satisfiable in the set of successors of w ,
then Σ is satisfiable in the set of successors of w

Let τ be a modal similarity type, and let \mathfrak{M} be a τ -model. \mathfrak{M} is called **m -saturated** if for every state w of \mathfrak{M} and every (n -ary) modal operator $\Delta \in \tau$ and sequence $\Sigma_1, \dots, \Sigma_n$ of sets of modal formulas, we have the following:

If for every sequence of finite subsets $\Delta_1 \subset \Sigma_1, \dots, \Delta_n \subseteq \Sigma_n$, there are
states v_1, \dots, v_n s.t. $Rwv_1 \dots v_n$ and $v_1 \Vdash \Delta_1, \dots, v_n \Vdash \Delta_n$,
then there are states v_1, \dots, v_n in \mathfrak{M} s.t. $Rwv_1 \dots v_n$ and $v_1 \Vdash \Sigma_1, \dots, v_n \Vdash \Sigma_n$

Proposition 2.37. *Let τ be a modal similarity type. Then the class of m -saturated τ -models has the Hennessy-Milner property*

Proof. Let $\mathfrak{M} = (W, R, V)$ and $\mathfrak{M}' = (W', R', V')$ be two m -saturated models.

Assume that $w, v \in W$ and $w' \in W'$ are s.t. Rwv and $w \leftrightarrow w'$. Let Σ be the set of formulas true at v . It is clear that for every finite subset Δ of Σ we have $\mathfrak{M}, v \Vdash \bigwedge \Delta$, hence $\mathfrak{M}, w \Vdash \Diamond \bigwedge \Delta$. As $w \leftrightarrow w'$, it follows that $\mathfrak{M}', w' \Vdash \Diamond \bigwedge \Delta$, so w' has an R' -successor v_Δ s.t. $\mathfrak{M}', v_\Delta \Vdash \bigwedge \Delta$. In other words, Σ is finitely satisfiable in the set of successors of w' ; but then, by m -saturation, Σ itself is satisfiable in a successor v' of w' . Thus $v \leftrightarrow v'$ \square

2.5.2 Ultrafilter extensions

Definition 2.38 (Filters and Ultrafilters). Let W be a non-empty set. A **filter** F **over** W is a set $F \subseteq \mathcal{P}(W)$ s.t.

1. $W \in F$
2. If $X, Y \in F$, then $X \cap Y \in F$
3. If $X \in F$ and $X \subseteq Z \subseteq W$, then $Z \in F$

An **ultrafilter over** W is a proper filter s.t. for all $X \in \mathcal{P}(W)$, $X \in U$ iff $(W \setminus X) \notin U$

Definition 2.39. Let W be a non-empty set, and let E be a subset of $\mathcal{P}(W)$. By the **filter generated by** E we mean the intersection F of the collection of all filters over W which include E

$$F = \bigcap \{G \mid E \subseteq G \text{ and } G \text{ is a filter over } W\}$$

E has the **finite intersection property** if the intersection of any finite number of elements of E is non-empty

Lemma 2.40 (Zorn's Lemma). *Whenever $<$ is a strict partial order of a set A satisfying for all chains $C \subseteq A$ there is some $b \in A$ s.t. $x \leq b$ for all $x \in C$ then for all $a \in A$, there is a maximal $b \in A$ with $b \geq a$*

Theorem 2.41 (Ultrafilter Theorem). *Fix a non-empty set W . Any proper filter over W can be extended to an ultrafilter over W . As a corollary, any subset of $\mathcal{P}(W)$ with the finite intersection property can be extended to an ultrafilter over W*

Definition 2.42. Let W be a non-empty set. Given an element $w \in W$, the **principal ultrafilter** π_w generated by w is the filter generated by the singleton set $\{w\}$

Suppose U is an ultrafilter over a non-empty set I , and that for each $i \in I$, A_i is a non-empty set. Let $C = \prod_{i \in I} A_i$. That is, C is the set of all functions f with domain I s.t. for each $i \in I$, $f(i) \in A_i$. For two functions $f, g \in C$ we say that f and g are **U -equivalent** ($f \sim_U g$) if $\{i \in I \mid f(i) = g(i)\} \in U$

Proposition 2.43. *The relation \sim_U is an equivalence relation on the set C*

Proof. Suppose $\{i \mid f(i) = g(i)\} \in U, \{i \mid g(i) = h(i)\} \in U$, then $\{i \mid f(i) = g(i) = h(i)\} = \{i \mid f(i) = g(i)\} \cap \{i \mid g(i) = h(i)\} \in U$. And $\{i \mid f(i) = g(i) = h(i)\} \subseteq \{i \mid f(i) = h(i)\}$ \square

Definition 2.44. Let f_U be the equivalence class of f modulo \sim_U , that is: $f_U = \{g \in C \mid g \sim_U f\}$. The **ultraproduct of the sets A_i modulo U** is the set of all equivalence classes of \sim_U . It is denoted by $\prod_U A_i$. So

$$\prod_U A_i = \{f_U \mid f \in \prod_{i \in I} A_i\}$$

Definition 2.45. Fix a first-order language \mathcal{L}^1 , and let $\mathfrak{A}_i (i \in I)$ be \mathcal{L}^1 -models. The **ultraproduct $\prod_U \mathfrak{A}_i$ of \mathfrak{A}_i modulo U** is the model described as follows:

1. The universe A_U is the set $\prod_U A_i$, where A_i is the universe of \mathfrak{A}_i
2. Let R be an n -place relation symbol, and R_i its interpretation in the model \mathfrak{A}_i . The relation R_U in $\prod_U \mathfrak{A}_i$ is given by

$$R_U f_U^1 \dots f_U^n \quad \text{iff} \quad \{i \in I \mid R_i f^1(i) \dots f^n(i)\} \in U$$

3. Let F be an n -place function symbol, and F_i its interpretation in \mathfrak{A}_i . The function F_U in $\prod_U \mathfrak{A}_i$ is given by

$$F_U(f_U^1, \dots, f_U^n) = \{(i, F_i(f^1(i), \dots, f^n(i))) \mid i \in I\}_U$$

4. Let c be a constant, and a_i its interpretation in \mathfrak{A}_i . Then c is interpreted by the element $c' \in \prod_U A_i$ where $c' = \{(i, a_i) \mid i \in I\}_U$

In the case where all the structures are the same, say $\mathfrak{A}_i = \mathfrak{A}$ for all i , we speak of the **ultrapower** of \mathfrak{A} modulo U , notation $\prod_U \mathfrak{A}$

Theorem 2.46 (Łoś's Theorem). *Let U be an ultrafilter over a non-empty set I . For each $i \in I$, let \mathfrak{A}_i be a model*

1. *For every term $t(x_1, \dots, x_n)$ and all elements f_U^1, \dots, f_U^n of $\mathfrak{B} = \prod_U \mathfrak{A}_i$ we have*

$$t^{\mathfrak{B}}[x_1 \mapsto f_U^1, \dots, x_n \mapsto f_U^n] = \{(i, t^{\mathfrak{A}_i}[f^1(i), \dots, f^n(i)]) \mid i \in I\}_U$$

2. *Given any first-order formula $\alpha(x_1, \dots, x_n)$ in \mathcal{L}_τ^1 and f_U^1, \dots, f_U^n in $\prod_U \mathfrak{A}_i$ we have*

$$\prod_U \mathfrak{A}_i \models \alpha[f_U^1, \dots, f_U^n] \quad \text{iff} \quad \{i \in I \mid \mathfrak{A}_i \models \alpha[f^1(i), \dots, f^n(i)]\} \in U$$

Proof. 1.

2. Induction on α . The atomic case holds by definition. Suppose that $\alpha \equiv \neg\beta(x_1, \dots, x_n)$, then

$$\begin{aligned} \prod_U \mathfrak{A}_i \models \alpha[f_U^1 \dots f_U^n] & \quad \text{iff} \quad \prod_U \mathfrak{A}_i \not\models \beta[f_U^1, \dots, f_U^n] \\ & \quad \text{iff} \quad \{i \in I \mid \mathfrak{A}_i \models \beta[f_U^1, \dots, f_U^n]\} \notin U \\ & \quad \text{iff} \quad \{i \in I \mid \mathfrak{A}_i \not\models \beta[f^1(i), \dots, f^n(i)]\} \in U \\ & \quad \text{iff} \quad \{i \in I \mid \mathfrak{A}_i \models \alpha[f^1(i), \dots, f^n(i)]\} \in U \end{aligned}$$

The second equivalence follows from the inductive hypothesis, and the third from the fact that U is an ultrafilter

Suppose that $\alpha(x_1, \dots, x_n) \equiv \exists x_0 \beta(x_0, \dots, x_n)$, then

$$\begin{aligned} \prod_U \mathfrak{A}_i \models \alpha[f_U^1, \dots, f_U^n] & \text{ iff } \exists f_U^0 \in \prod_U \mathfrak{A}_i, \prod_U \mathfrak{A}_i \models \beta[f_U^0, \dots, f_U^n] \\ & \text{ iff } \exists f_U^0 \in \prod_U \mathfrak{A}_i, \{i \in I \mid \mathfrak{A}_i \models \beta[f^0(i), \dots, f^n(i)]\} \in U \end{aligned} \quad (1)$$

As $\mathfrak{A}_i \models \beta[f^0(i), \dots, f^n(i)]$ implies $\mathfrak{A}_i \models \alpha[f^1(i), \dots, f^n(i)]$, which means

$$\{i \in I \mid \mathfrak{A}_i \models \beta[f^0(i), \dots, f^n(i)]\} \subseteq \{i \in I \mid \mathfrak{A}_i \models \alpha[f^1(i), \dots, f^n(i)]\}$$

Hence

$$\{i \in I \mid \mathfrak{A}_i \models \alpha[f^1(i), \dots, f^n(i)]\} \in U \quad (2)$$

Conversely, if (2) holds, then we can select a function $f^0 \in \prod_{i \in I} A_i$ s.t. (1) holds. So (1) is equivalent to (2)

□

Corollary 2.47. *Let $\prod_U \mathfrak{A}$ be an ultrapower of \mathfrak{A} . Then for all first-order sentences α , $\mathfrak{A} \models \alpha$ iff $\prod_U \mathfrak{A} \models \alpha$*

There is a natural embedding of a model \mathfrak{A} in each of its ultrapowers. Define the **diagonal mapping** d of \mathfrak{A} into $\prod_U \mathfrak{A}$ to be the function

$$\alpha \mapsto (f_\alpha)_U, \text{ where } f_\alpha(i) = a \text{ for all } i \in I$$

Corollary 2.48. *Let $\prod_U \mathfrak{A}$ be an ultrapower of \mathfrak{A} . Then the diagonal mapping of \mathfrak{A} into $\prod_U \mathfrak{A}$ is an elementary embedding*

Proof.

$$\begin{aligned} \prod_U \mathfrak{A} \models \alpha[d(a_1), \dots, d(a_n)] & \text{ iff } \{i \in I \mid \mathfrak{A} \models \alpha[a_1, \dots, a_n]\} \in U \\ & \text{ iff } \mathfrak{A} \models \alpha[a_1, \dots, a_n] \end{aligned}$$

□

$$V(\phi) = \{w \mid \mathfrak{M}, w \Vdash \phi\}$$

Definition 2.49. Given an $(n + 1)$ -ary relation R on a set W , we define the following two n -ary operations m_R and l_R on the power set $\mathcal{P}(W)$ of W :

$$\begin{aligned} m_R(X_1, \dots, X_n) &:= \{w \in W \mid \exists w_1, \dots, w_n (Rww_1 \dots w_n \bigwedge \forall i (w_i \in X_i))\} \\ l_R(X_1, \dots, X_n) &:= \{w \in W \mid \forall w_1, \dots, w_n (Rww_1 \dots w_n \rightarrow \exists i (w_i \in X_i))\} \\ m_R(V(\phi_1), \dots, V(\phi_n)) &:= V(\Delta(\phi_1, \dots, \phi_n)) \\ l_R(V(\phi_1), \dots, V(\phi_n)) &:= V(\nabla(\phi_1, \dots, \phi_n)) \end{aligned}$$

It follows that for any model $\mathfrak{M} = (W, R, V)$ we have

$$V(\Diamond\phi) = m_R(V(\phi)) \quad \text{and} \quad V(\Box\phi) = l_R(V(\phi))$$

Proposition 2.50. Let R be a relation of arity $n + 1$ on the set W . Then for every n -tuple X_1, \dots, X_n of subsets of W we have

$$l_R(X_1, \dots, X_n) = W \setminus m_R(W \setminus X_1, \dots, W \setminus X_n)$$

Proof. This is actually $\nabla = \neg\Delta\neg$

$$\begin{aligned} W \setminus m_R(W \setminus X_1, \dots, W \setminus X_n) &= \{w \mid \neg \exists w_1, \dots, w_n (Rww_1 \dots w_n \bigwedge \forall i (w_i \in W \setminus X_i))\} \\ &= \{\forall w_1, \dots, w_n (\neg Rww_1 \dots w_n \bigvee \neg \forall i (w_i \in W \setminus X_i))\} \\ &= \{\forall w_1, \dots, w_n (Rww_1 \dots w_n \rightarrow \exists i (w_i \notin W \setminus X_i))\} \\ &= l_R(X_1, \dots, X_n) \end{aligned}$$

□

Definition 2.51 (Ultrafilter Extension). Let τ be a modal similarity type, and $\mathfrak{F} = (W, R_\Delta)_{\Delta \in \tau}$ is a τ -frame. The **ultrafilter extension** $\text{ue}\mathfrak{F}$ of \mathfrak{F} is defined as the frame $(Uf(W), R_\Delta^{ue})_{\Delta \in \tau}$. Here $Uf(W)$ is the set of ultrafilters over W and $R_\Delta^{ue} u_0 u_1 \dots u_n$ holds for a tuple u_0, \dots, u_n of ultrafilters over W if we have that $m_{R_\Delta}(X_1, \dots, X_n) \in u_0$ whenever $X_i \in u_i$ for all i with $1 \leq i \leq n$.

The **ultrafilter extension** of a τ -model $\mathfrak{M} = (\mathfrak{F}, V)$ is the model $\text{ue}\mathfrak{M} = (\text{ue}\mathfrak{F}, V^{ue})$ where $V^{ue}(p_i)$ is the set of ultrafilters of which $V(p_i)$ is a member

Any subset of a frame can be viewed as a **proposition**. A filter over the universe of the frame can thus be seen as a **theory**, in fact as a logically closed theory, since filters are both closed under intersection (conjunction) and upward closed (entailment). Viewed this way, a proper filter is a **consistent** theory, or **state of affairs**, for it does not contain the empty set (falsum). Finally an ultrafilter is a **complete** theory.

In a given frame \mathfrak{F} not every state of affairs needs to 'realized', in the sense that there is a state satisfying all and only the propositions belonging to the state of affairs; only the states of affairs that correspond to the **principal** ultrafilters are realized. We build $ue\mathfrak{F}$ by adding every state of affairs for \mathfrak{F} as a new element of the domain - that is, $ue\mathfrak{F}$ realizes every proposition in \mathfrak{F}

Stipulate that $R_{\Delta}^{ue}u_0u_1 \dots u_n$ if u_0 'sees' the n -tuple u_1, \dots, u_n . That is, whenever X_1, \dots, X_n are propositions of u_1, \dots, u_n respectively, then u_0 'sees' this combination: that is, the proposition $m_{R_{\Delta}}(X_1, \dots, X_n)$ is a member of u_0 .

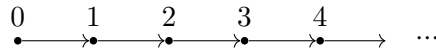
Principal ultrafilters over W plays a special role. By identifying a state w of a frame \mathfrak{F} with the principal ultrafilter $\pi_w = \{X \subseteq W \mid w \in X\}$, it is easily seen that any frame \mathfrak{F} is (isomorphic to) a **submodel** (but in general not a **generated** submodel) of its ultrafilter extension. For we have the following equivalences

$$\begin{aligned} R w v & \text{ iff } w \in m_R(X) \text{ for all } X \subseteq W \text{ s.t. } v \in X \\ & \text{ iff } m_R(X) \in \pi_w \text{ for all } X \subseteq W \text{ s.t. } X \in \pi_v \\ & \text{ iff } R^{ue} \pi_w \pi_v \end{aligned}$$

since

$$R w v \text{ iff } \forall X \subseteq W (v \in X \rightarrow w \in m_R(X))$$

Example 2.4. Consider the frame $\mathfrak{N} = (\mathbb{N}, <)$



What is the ultrafilter extension of \mathfrak{N} ? There are two kinds of ultrafilter over an infinite set: the principal ultrafilters that are in one-to-one correspondence with the points of the set, and the non-principal ones which contain all co-finite sets and only infinite sets, cf Exercise 2.5.1. The principal ultrafilters form an isomorphic copy of the frame \mathfrak{N} inside $ue\mathfrak{N}$. For any pair u, u' of ultrafilters, if u' is non-principal, then $R^{ue}uu'$. To see this, let $X \in u'$. As X is infinite, for any $n \in \mathbb{N}$ there is an m s.t. $n < m$ and $m \in X$. This shows that $m_{<}(X) = \mathbb{N}$. But \mathbb{N} is an element of every ultrafilter

This shows that the ultrafilter extension of \mathfrak{N} consists of a copy of \mathfrak{N} followed by an uncountable cluster consisting of all the non-principal ultrafilters

Proposition 2.52. Let τ be a modal similarity type, and \mathfrak{M} a τ -model. Then for any formula ϕ and any ultrafilter u over W , $V(\phi) \in u$ iff $ue\mathfrak{M}, u \models \phi$. Hence for every state w of \mathfrak{M} we have $w \rightsquigarrow \pi_w$

Proof. The second claim of the proposition is immediate from the first one by the observation that $w \Vdash \phi$ iff $w \in V(\phi)$ iff $V(\phi) \in \pi_w$

Induction on ϕ . The basic case is immediate from the definition of V^{ue} . Suppose ϕ is of the form $\neg\psi$, then

$$\begin{aligned} V(\neg\psi) \in u & \quad \text{iff} \quad W \setminus V(\psi) \in u \\ & \quad \text{iff} \quad V(\psi) \notin u \\ & \quad \text{iff} \quad u \notin \mathfrak{M}, u \not\models \psi \quad \text{IH} \\ & \quad \text{iff} \quad u \models \neg\psi \end{aligned}$$

Now consider the case where ϕ is of the form $\diamond\psi$. Assume first that $u \models \diamond\psi$. Then there is an ultrafilter u' s.t. $R^{ue}uu'$ and $u' \models \psi$. The induction hypothesis implies that $V(\psi) \in u'$, so by the definition of R^{ue} , $m_R(V(\psi)) \in u$. Now the result follows immediately from the observation that $m_R(V(\psi)) = V(\diamond\psi)$

Assume that $V(\diamond\psi) \in u$. We have to find an ultrafilter u' s.t. $V(\psi) \in u'$ and $R^{ue}uu'$. The latter constraint reduces to the condition that $m_R(X) \in u$ whenever $X \in u'$, or equivalently (see Exercise 2.5.2)

$$u'_0 := \{Y \mid l_R(Y) \in u\} \subseteq u'$$

We will first show that u'_0 is closed under intersection. Let $Y, Z \in u'_0$. By definition, $l_R(Y)$ and $l_R(Z)$ are in u . But then $l_R(Y \cap Z) \in u$ as $l_R(Y \cap Z) = l_R(Y) \cap l_R(Z)$. This proves that $Y \cap Z \in u'_0$

Next we make sure that for any $Y \in u'_0$, $Y \cap V(\psi) \neq \emptyset$. Let Y be an arbitrary element of u'_0 , then by definition of u'_0 , $l_R(Y) \in u$. As u is closed under intersection and does not contain the empty set, there must be an element $x \in l_R(Y) \cap V(\diamond\psi)$. But then x must have a successor y in $V(\psi)$. Finally, $x \in l_R(Y)$ implies $y \in Y$

From the fact that u'_0 is closed under intersection, and the fact that for any $Y \in u'_0$, $Y \cap V(\psi) \neq \emptyset$, it follows that the set $u'_0 \cup \{V(\psi)\}$ has the finite intersection property. So the Ultrafilter Theorem provides us with an ultrafilter u' s.t. $u'_0 \cup \{V(\psi)\} \subseteq u'$. This ultrafilter u' has the desired properties: it is clearly a successor of u , and the fact the $u \models \psi$ follows from $V(\psi) \in u'$ and the induction hypothesis \square

Example 2.5. Our new invariance result can be used to compare the relative expressive power of modal languages. Consider the modal constant \mathcal{O} whose truth definition in a model for the basic modal language is

$$\mathfrak{M}, w \Vdash \mathcal{O} \quad \text{iff} \quad \mathfrak{M} \models Rxx[v] \text{ for some } v \text{ in } \mathfrak{M}$$

Comparing the pictures of the frame $(\mathbb{N}, <)$ and its ultrafilter extension given in Example 2.4 . The former is loop-free but the latter contains uncountably many loops

Proposition 2.53. *Let τ be a modal similarity type, and let \mathfrak{M} be a τ -model. Then $ue\mathfrak{M}$ is m -saturated*

Proof. Let $\mathfrak{M} = (W, R, V)$ be a model. Consider an ultrafilter u over W , and a set Σ of modal formulas which is finitely satisfiable in the set of successors of u . We have to find an ultrafilter u' s.t. $R^{ue}uu'$ and $ue\mathfrak{M}, u' \models \Sigma$. Define

$$\Delta = \{V(\phi) \mid \phi \in \Sigma'\} \cup \{Y \mid l_R(Y) \in u\}$$

where Σ' is the set of (finite) conjunctions of formulas in Σ . We claim that the set Δ has the finite intersection property. Since both $\{V(\phi) \mid \phi \in \Sigma'\}$ and $\{Y \mid l_R(Y) \in u\}$ are closed under taking intersections, it suffices to prove that for an arbitrary $\phi \in \Sigma'$ and an arbitrary set $Y \subseteq W$ for which $l_R(Y) \in u$, we have $V(\phi) \cap Y \neq \emptyset$. but if $\phi \in \Sigma'$, then by assumption, there is a successor u'' of u s.t. $ue\mathfrak{M}, u'' \models \phi$, or in other words, $V(\phi) \in u''$. Then $l_R(Y) \in u$ implies $Y \in u''$ by Exercise 2.5.2 . Hence $V(\phi) \cap Y$ is an element of the ultrafilter u'' and therefore cannot be identical to the empty set.

It follows by the Ultrafilter Theorem that Δ can be extended to an ultrafilter u' . Clearly u' is the required successor \square

Theorem 2.54. *Let τ be a modal similarity type, and let \mathfrak{M} and \mathfrak{M}' be τ -models, and w, w' two states in \mathfrak{M} and \mathfrak{M}' respectively. Then*

$$\mathfrak{M}, w \rightsquigarrow \mathfrak{M}', w' \quad \text{iff} \quad ue\mathfrak{M}, \pi_w \cong ue\mathfrak{M}', \pi_{w'}$$

Proof. From Propositions 2.52, 2.53 and 2.37 \square

Exercise 2.5.1. Let W be an infinite set. Recall that $X \subseteq W$ is **co-finite** if $W \setminus X$ is finite

1. Prove that the collection of co-finite subsets of W has the finite intersection property
2. Show that there are ultrafilters over W that do not contain any finite set
3. Prove that an ultrafilter is non-principal iff it contains only infinite sets iff it contains all co-finite sets
4. Prove that any ultrafilter over W has uncountably many elements

Proof. Suppose $U = \{X \subseteq W \mid X \text{ is cofinite}\}$

1. For any $A, B \in U$, if $A \cap B = \emptyset$, $A \subset \overline{B}$. But A is infinite and \overline{B} is finite, this can't happen. Hence $A \cap B \neq \emptyset$
2. U can be extended to a ultrafilter \mathcal{U} . If A is finite, then $\overline{A} \in U \subseteq \mathcal{U}$. Hence \mathcal{U} does not contain any finite set.
3. $1 \rightarrow 2$. If an ultrafilter contains a finite set. Then its a principal ultrafilter generated on the intersection of all finite sets.
- $2 \rightarrow 3$ and $3 \rightarrow 1$ are obvious.
4. Half of the $\mathcal{P}(W)$ belongs to the ultrafilter and $\mathcal{P}(W)$ is uncountable

□

Exercise 2.5.2. Given a model $\mathfrak{M} = (W, R, V)$ and two ultrafilters u and v over W , show that $R^{ue}uv$ iff $\{Y \mid l_R(Y) \in u\} \subseteq v$

Proof.

$$\begin{aligned}
 R^{ue}uv &\Leftrightarrow X \in v \rightarrow m_R(X) \in u \\
 &\Leftrightarrow \neg m_R(X) \in u \rightarrow \neg X \in v \\
 &\Leftrightarrow W - m_R(X) \in u \rightarrow W - X \in v \\
 &\Leftrightarrow l_R(W - X) \in u \rightarrow W - X \in v \\
 &\text{(Since } m_R(X) = W - l_R(W - X)\text{)} \\
 &\Leftrightarrow \{Y \mid l_R(Y) \in u\} \subseteq v
 \end{aligned}$$

□

2.6 Characterization and Definability

2.6.1 The van Benthem Characterization Theorem

Let $\Gamma(x)$ be a set of first-order formulas in which a single individual variable may occur free - such a set of formulas is called a **type**. A first-order model \mathfrak{M} **realizes** $\Gamma(x)$ if there is an element w in \mathfrak{M} s.t. for all $\gamma \in \Gamma$, $\mathfrak{M} \models \gamma[w]$

Let \mathfrak{M} be a model for a given first-order language \mathcal{L}^1 with domain W . For a subset $A \subset W$, $\mathcal{L}^1[A]$ is the language obtained by extending \mathcal{L}^1 with new constant \underline{a} for all elements $a \in A$. \mathfrak{M}_A is the expansion of \mathfrak{M} to a structure for $\mathcal{L}^1[A]$ in which each \underline{a} is interpreted as a

Assume that A is of size at most α . Assume that $\alpha = 3$ and $A = \{\alpha_1, \alpha_2\}$. Let $\Gamma(\underline{a}_1, \underline{a}_2, x)$ be a type of the language $\mathcal{L}^1[A]$; $\Gamma(\underline{a}_1, \underline{a}_2, x)$ is consistent with the first-order theory of \mathfrak{M}_A iff $\Gamma(\underline{a}_1, \underline{a}_2, x)$ is finitely realizable in \mathfrak{M}_A . So for this particular set $\Gamma(\underline{a}_1, \underline{a}_2, x)$, 3-saturation of \mathfrak{M} means that if $\Gamma(\underline{a}_1, \underline{a}_2, x)$ is finitely realizable in \mathfrak{M}_A , then $\Gamma(\underline{a}_1, \underline{a}_2, x)$ is realizable in \mathfrak{M}_A .

Or consider a formula $\gamma(\underline{a}_1, \underline{a}_2, x)$ and let $\gamma(x_1, x_2, x)$ be the formula with the fresh variables x_1 and x_2 replacing each occurrence in γ of \underline{a}_1 and \underline{a}_2 respectively. Then we have the following equivalence

$$\mathfrak{M}_A \text{ realizes } \{\gamma(\underline{a}_1, \underline{a}_2, x)\} \text{ iff there is a } b \text{ s.t. } \mathfrak{M} \models \gamma(x_1, x_2, x)[a_1, a_2, b]$$

So a model is α -saturated iff the following holds for every $n < \alpha$ and every set Γ of formulas of the form $\gamma(x_1, \dots, x_n, x)$

If (a_1, \dots, a_n) is an n -tuple s.t. for every finite $\Delta \subseteq \Gamma$ there is a b_Δ s.t.

$$\mathfrak{M} \models \gamma(x_1, \dots, x_n, x)[a_1, \dots, a_n, b_\Delta] \text{ for every } \gamma \in \Delta$$

then we have that there is a b s.t. $\mathfrak{M} \models \gamma(x_1, \dots, x_n, x)[a_1, \dots, a_n, b]$ for every $\gamma \in \Gamma$

Definition 2.55. Let α be a natural number, or ω . A model \mathfrak{M} is α -saturated if for every subset $A \subseteq W$ of size less than α , the expansion \mathfrak{M}_A realizes every set $\Gamma(x)$ of $\mathcal{L}^1[A]$ -formulas (with only x occurring free) that is *consistent* (a proof-theoretic notion, only finite deductions, hence this definition is consistent the definition above) with the first-order theory of \mathfrak{M}_A . An ω -saturated model is usually called **countably saturated**

Example 2.6. 1. Every finite model is countably saturated. For if \mathfrak{M} is finite, and $\Gamma(x)$ is a set of first-order formulas consistent with the first-order theory of \mathfrak{M} , there exists a model \mathfrak{N} that is elementarily equivalent to \mathfrak{M} and that realizes $\Gamma(x)$. But as \mathfrak{M} and \mathfrak{N} are finite, elementary equivalence implies isomorphism (proof), and hence $\Gamma(x)$ is realized in \mathfrak{M}

2. The ordering of the rational numbers $(\mathbb{Q}, <)$ is countably saturated as well. The relevant first-order language \mathcal{L}^1 has $<$ and $=$. Take a subset A of \mathbb{Q} and let $\Gamma(x)$ be a set of formulas in the resulting expansion $\mathcal{L}^1[A]$ of the first-order language that is consistent with the theory of $(\mathbb{Q}, <, a)_{a \in A}$. Then there exists a model \mathfrak{N} of the theory of $(\mathbb{Q}, <, a)_{a \in A}$ that realizes $\Gamma(x)$. \star Now take a countable elementary submodel \mathfrak{N}' of \mathfrak{N} that contains at least one object realizing $\Gamma(x)$. Then \mathfrak{N}' is a countable dense linear ordering without endpoints, and hence the ordering of \mathfrak{N}' is isomorphic to $(\mathbb{Q}, <)$.

Theorem 2.56. *Let τ be a modal similarity type. Any countably saturated τ -model is m -saturated. It follows that the class of countably saturated τ -models has the Hennessy-Milner property*

Proof. Assume that $\mathfrak{M} = (W, R, V)$ viewed as a first-order model, is countably saturated. Let a be a state in W , and consider a set Σ of modal formulas which is finite satisfiable in the successor set of a . Define Σ' to be the set

$$\Sigma' = \{Rax\} \cup ST_x(\Sigma)$$

where $ST_x(\Sigma) = \{ST_x(\phi) \mid \phi \in \Sigma\}$. Σ' is consistent with the first-order theory of \mathfrak{M}_a : \mathfrak{M}_a realizes every finite subset of Σ' , namely in some successor of a . So by the countable saturation of \mathfrak{M} , Σ' is realized in some state b . By $\mathfrak{M}_A \models Rax[b]$ it follows that b is a successor of a . Then, by Proposition 2.32 and the fact that $\mathfrak{M}_a \models ST_x(\phi)[b]$ for all $\phi \in \Sigma$, it follows that $\mathfrak{M}, b \models \Sigma$. Thus Σ is satisfiable in a successor of a \square

Lemma 2.57 (Detour Lemma). *Let τ be a modal similarity type, and let \mathfrak{M} and \mathfrak{N} be two models, and w and v states in \mathfrak{M} and \mathfrak{N} , respectively. Then the following are equivalent:*

1. *For all modal formulas ϕ : $\mathfrak{M}, w \models \phi$ iff $\mathfrak{N}, v \models \phi$*
2. *There exists a bisimulation $Z : \text{ue}\mathfrak{M}, \pi_w \rightleftharpoons \text{ue}\mathfrak{N}, \pi_v$*
3. *There exist countably saturated models $\mathfrak{M}^*, w^*, \mathfrak{N}^*, v^*$ and elementary embeddings $f : \mathfrak{M} \preceq \mathfrak{M}^*$ and $g : \mathfrak{N} \preceq \mathfrak{N}^*$ s.t.*
 - (a) $f(w) = w^*$ and $g(v) = v^*$
 - (b) $\mathfrak{M}^*, w^* \rightleftharpoons \mathfrak{N}^*, v^*$

Definition 2.58. A first-order formula $\alpha(x)$ in \mathcal{L}_τ^1 is **invariant for bisimulations** if for all models \mathfrak{M} and \mathfrak{N} , and all states w in \mathfrak{M} , v in \mathfrak{N} , and all bisimulations Z between \mathfrak{M} and \mathfrak{N} s.t. wZv , we have $\mathfrak{M} \models \alpha(x)[w]$ iff $\mathfrak{N} \models \alpha(x)[v]$

Theorem 2.59 (van Benthem Characterization Theorem). *Let $\alpha(x)$ be a first-order formula in \mathcal{L}_τ^1 . Then $\alpha(x)$ is invariant for bisimulations iff it is equivalent to the standard translation of a modal τ -formula*

Proof. Assume $\alpha(x)$ is invariant for bisimulations and consider the set of modal consequences of α :

$$MOC(\alpha) = \{ST_x(\phi) \mid \phi \text{ is a modal formula, and } \alpha(x) \models ST_x(\phi)\}$$

Our first claim is that if $MOC(\alpha) \models \alpha(x)$, then α is equivalent to the translation of a modal formula. Assume $MOC(\alpha) \models \alpha(x)$, then by the Compactness Theorem for first-order logic, for some finite subset $X \subseteq MOC(\alpha)$, we have $X \models \alpha(x)$. So $\models \bigwedge X \rightarrow \alpha(x)$. Trivially $\models \alpha(x) \rightarrow \bigwedge X$, thus $\models \alpha(x) \leftrightarrow \bigwedge X$. And as every $\beta \in X$ is the translation of a modal formula, so is $\bigwedge X$.

So it suffices to show that $MOC(\alpha) \models \alpha(x)$. Assume $\mathfrak{M} \models MOC(\alpha)[w]$; we need to show that $\mathfrak{M} \models \alpha(x)[w]$. Let

$$T(x) = \{ST_x(\phi) \mid \mathfrak{M} \models ST_x(\phi)[w]\}$$

We claim that $T(x) \cup \{\alpha(x)\}$ is consistent. Assume that $T(x) \cup \{\alpha(x)\}$ is inconsistent. Then by compactness, for some finite subset $T_0(x) \subset T(x)$ we have $\models \alpha(x) \rightarrow \neg \bigwedge T_0(x)$. Hence $\neg \bigwedge T_0(x) \in MOC(\alpha)$. But this implies $\mathfrak{M} \models \neg \bigwedge T_0(x)[w]$, a contradiction.

Let \mathfrak{N}, v be s.t. $\mathfrak{N} \models T(x) \cup \{\alpha(x)\}[v]$. Observe that w and v are modally equivalent: $\mathfrak{M}, w \Vdash \phi$ implies $ST_x(\phi) \in T(x)$, which implies $\mathfrak{N}, v \Vdash \phi$; and likewise, if $\mathfrak{M}, w \nVdash \phi$ then $\mathfrak{M}, w \Vdash \neg \phi$ and $\mathfrak{N}, v \Vdash \neg \phi$.

We can use the Detour Lemma and make a detour through a Hennessy-Milner class where modal equivalence and bisimilarity do coincide.

$$\begin{array}{ccc} \mathfrak{M}, w & & \mathfrak{N}, v \\ \downarrow \preceq & & \downarrow \preceq \\ \mathfrak{M}^*, w^* & \xrightarrow{\cong} & \mathfrak{N}^*, v^* \end{array}$$

$\mathfrak{M} \models \alpha(x)[w]$ implies $\mathfrak{N}^* \models \alpha(x)[v^*]$. As $\alpha(x)$ is invariant for bisimulations, we get $\mathfrak{M}^* \models \alpha(x)[w^*]$. By invariance under elementary embeddings, we have $\mathfrak{M} \models \alpha(x)[w]$. \square

2.6.2 Ultraproducts

Suppose $I \neq \emptyset$, U is an ultrafilter over I .

Definition 2.60 (Ultraproducts of Sets). Let f_U be the equivalence class of f modulo \sim_U , that is: $f_U = \{g \in C \mid g \sim_U f\}$. The **ultraproduct of W_i modulo U** , denoted as $\prod_U W_i$ is the set of all equivalence classes of \sim_U . So

$$\prod_U W_i = \{f_U \mid f \in \prod_{i \in I} W_i\}$$

In case $W_i = W$, the ultraproduct is called the **ultrapower of W modulo U** , and written $\prod_U W$.

Definition 2.61 (Ultraproduct of Models). Fix a modal similarity type τ , and let $\mathfrak{M}_i (i \in I)$ be τ -models. The **ultraproduct** $\prod_U \mathfrak{M}_i$ of \mathfrak{M}_i modulo U is the model described as follows

1. The universe W_U of $\prod_U \mathfrak{M}_i$ is the set $\prod_U W_i$
2. Let V_i be the valuation of \mathfrak{M}_i . Then the valuation V_U of $\prod_U \mathfrak{M}_i$ is defined by

$$f_U \in V_U(p) \quad \text{iff} \quad \{i \in I \mid f(i) \in V_i(p)\} \in U$$

3. Let \triangle be a modal operator in τ , and $R_{\triangle i}$ its associated relation in the model \mathfrak{M}_i . The relation $R_{\triangle U}$ in $\prod_U \mathfrak{M}_i$ is given by

$$R_{\triangle U} f_U^1 \dots f_U^{n+1} \quad \text{iff} \quad \{i \in I \mid R_{\triangle i} f^1(i) \dots f^{n+1}(i)\} \in U$$

In particular,

$$R_{\diamond U} f_U g_U \quad \text{iff} \quad \{i \in I \mid R_{\diamond i} f(i)g(i)\} \in U$$

Proposition 2.62. Let $\prod_U \mathfrak{M}$ be an ultrapower of \mathfrak{M} . Then for all modal formulas ϕ we have $\mathfrak{M}, w \Vdash \phi$ iff $\prod_U \mathfrak{M}, (f_w)_U \Vdash \phi$, where f_w is the constant function s.t. $f_w(i) = w$ for all $i \in I$

Proof. 1. $\phi = p$

$$\begin{aligned} \mathfrak{M}, w \Vdash \phi &\Leftrightarrow w \in V(\phi) \\ &\Leftrightarrow \{i \in I \mid f_w(i) \in V(p)\} = I \in U \\ &\Leftrightarrow \prod_U \mathfrak{M}, (f_w)_U \Vdash \phi \end{aligned}$$

2. $\phi = \diamond\psi$

$$R w v \Leftrightarrow \{i \in I \mid R_{\diamond i} f_w(i) f_v(i)\} = I \in U \Leftrightarrow R_{\diamond U} f_w g_v$$

□

An ultrafilter is **countably incomplete** if it is not closed under countably intersections

Example 2.7. Consider the set of natural numbers \mathbb{N} . Let U be an ultrafilter over \mathbb{N} that does not contain any singletons $\{u\}$. Then for all n , $(\mathbb{N} \setminus \{n\}) \in U$. But

$$\emptyset = \bigcap_{n \in \mathbb{N}} (\mathbb{N} \setminus \{n\}) \notin U$$

So U is countably incomplete

Lemma 2.63. *Let \mathcal{L} be a countable first-order language, U a countably incomplete ultrafilter over a non-empty set I , and \mathfrak{M} an \mathcal{L} -model. The ultrapower $\prod_U \mathfrak{M}$ is countably saturated*

Theorem 2.64. *Let τ be a modal similarity type, and let \mathfrak{M} and \mathfrak{N} be τ -models, and w and v states in \mathfrak{M} and \mathfrak{N} respectively. Then the following are equivalent*

1. *For all modal formulas ϕ : $\mathfrak{M}, w \Vdash \phi$ iff $\mathfrak{N}, v \Vdash \phi$*
2. *There exists ultrapowers $\prod_U \mathfrak{M}$ and $\prod_U \mathfrak{N}$ as well as a bisimulation $Z : \prod_U \mathfrak{M}, (f_w)_U \rightleftharpoons \prod_U \mathfrak{N}, (f_v)_U$ linking $(f_w)_U$ and $(f_v)_U$, where $f_w(f_v)$ is the constant function mapping every index to $w(v)$*

Proof. $2 \rightarrow 1$. $\mathfrak{M}, w \Vdash \phi$ iff $\prod_U \mathfrak{M}, (f_w)_U \Vdash \phi$ iff $\prod_U \mathfrak{N}, (f_v)_U \Vdash \phi$ iff $\mathfrak{N}, v \Vdash \phi$
 $1 \rightarrow 2$. Take \mathbb{N} as our index set, and let U be a countably incomplete ultrafilter over \mathbb{N} . By Lemma 2.63 the ultrapowers $\prod_U \mathfrak{M}$ and $\prod_U \mathfrak{N}$ are countably saturated. Now $(f_w)_U$ and $(f_v)_U$ are modally equivalent. Next apply Theorem 2.56: as $(f_w)_U$ and $(f_v)_U$ are modally equivalent and $\prod_U \mathfrak{M}$ and $\prod_U \mathfrak{N}$ are countably saturated, there exists a bisimulation \square

2.6.3 Definability

Given a modal similarity type τ , a pointed model is a pair (\mathfrak{M}, w) where \mathfrak{M} is a τ -model and w is a state of \mathfrak{M} . A class of pointed models K is said to be **closed under bisimulations** if $(\mathfrak{M}, w) \in K$ and $\mathfrak{M}, w \rightleftharpoons \mathfrak{N}, v$ implies $(\mathfrak{N}, v) \in K$. K is **closed under ultraproducts** if any ultraproducts $\prod_U (\mathfrak{M}_i, w_i)$ of a family of pointed models (\mathfrak{M}_i, w_i) in K belongs to K . If K is a class of pointed τ -models, \bar{K} denotes the complement of K within the class of all pointed τ -models. K is **definable by a set of modal formulas** if there is a set of modal formulas Γ s.t. for any pointed model (\mathfrak{M}, w) we have $(\mathfrak{M}, w) \in K$ iff for all $\gamma \in \Gamma$, $\mathfrak{M}, w \Vdash \gamma$. K is definable by a single modal formula iff it is definable by a singleton set

By Theorem 2.12 definable classes of pointed models must be closed under bisimulations, and by Proposition 2.32 and Corollary 2.47 they must be closed under ultraproducts as well.

Theorem 2.65. *Let τ be a modal similarity type, and K a class of pointed τ -models. Then the following are equivalent:*

1. *K is definable by a set of modal formulas*
2. *K is closed under bisimulations and ultraproducts, and \bar{K} is closed under ultrapowers*

Proof. Assume K and \bar{K} satisfy the stated closure conditions. Observe that \bar{K} is closed under bisimulations as K is. Define

$$T = \{\phi \mid \forall (\mathfrak{M}, w) \in K : \mathfrak{M}, w \Vdash \phi\}$$

We will show that T defines the class of K .

Assume $\mathfrak{M}, w \Vdash T$. Define $\Sigma = \{\phi \mid \mathfrak{M}, w \Vdash \phi\}$. It is obvious that Σ is finitely satisfiable in K ; for suppose that the set $\{\sigma_1, \dots, \sigma_n\} \subseteq \Sigma$ is not satisfiable in K . Then the formula $\neg(\sigma_1 \wedge \dots \wedge \sigma_n)$ would be true on all pointed models in K , so it would belong to T , yet be false in \mathfrak{M}, w . But then the following claim shows that Σ is satisfiable in the ultraproduct of pointed models

Claim 1 . Let Σ be a set of modal formulas, and K a class of pointed models in which Σ is finitely satisfiable. Then Σ is satisfiable in some ultraproduct of models in K

Proof of Claim. Define index set I as the collection of all finite subsets of Σ

$$I = \{\Sigma_0 \subseteq \Sigma \mid \Sigma_0 \text{ is finite}\}$$

By assumption, for each $i \in I$ there is a pointed model (\mathfrak{N}_i, v_i) in K s.t. $\mathfrak{N}_i, v_i \Vdash i$. We now construct an ultrafilter U over I s.t. the ultraproduct $\prod_U \mathfrak{N}_i$ has a state f_U with $\prod_U \mathfrak{N}_i, f_U \Vdash \Sigma$

For each $\sigma \in \Sigma$, let $\hat{\sigma}$ be the set of all $i \in I$ s.t. $\sigma \in i$. Then the set $E = \{\hat{\sigma} \mid \sigma \in \Sigma\}$ has the finite intersection property because

$$\{\sigma_1, \dots, \sigma_n\} \in \hat{\sigma}_1 \cap \dots \cap \hat{\sigma}_n$$

So E can be extended to an ultrafilter U over I . This defines $\prod_U \mathfrak{N}_i$; for the definition of f_U , let W_i denote the universe of the model \mathfrak{N}_i and consider the function $f \in \prod_{i \in I} W_i$ s.t. $f(i) = v_i$. It is left to prove that

$$\prod_U \mathfrak{N}_i, f_U \Vdash \Sigma$$

Observe that for $i \in \hat{\sigma}$ we have $\sigma \in i$ and so $\mathfrak{N}_i, v_i \Vdash \sigma$. Therefore for each $\sigma \in \Sigma$

$$\{i \in I \mid \mathfrak{N}_i, v_i \Vdash \sigma\} \supseteq \hat{\sigma} \quad \text{and} \quad \hat{\sigma} \in U$$

since $\sigma \in i$ implies $\mathfrak{N}_i, v_i \Vdash \sigma$. It follows that $\{i \in I \mid \mathfrak{N}_i, v_i \Vdash \sigma\} \in U$, so by Theorem 2.46 $\prod_U \mathfrak{N}_i, f_U \Vdash \sigma$. Hence $\prod_U \mathfrak{N}_i, f_U \Vdash \Sigma$

It follows from Claim 1 and the closure of K under taking ultraproducts that Σ is satisfiable in some pointed model $(\mathfrak{N}, v) \in K$. But $\mathfrak{N}, v \Vdash \Sigma$ implies

that v and the state w from our original pointed model (\mathfrak{M}, w) are modally equivalent. So by Theorem 2.64 there exists an ultrafilter U' s.t.

$$\prod_{U'} (\mathfrak{N}, v), (f_v)_U \Leftrightarrow \prod_{U'} (\mathfrak{M}, w), (f_w)_U$$

By closure under ultraproducts, the pointed model $(\prod_{U'} (\mathfrak{N}, v), (f_v)_U)$ belongs to K . Hence by closure under bisimulations, $(\prod_{U'} (\mathfrak{M}, w), (f_w)_U)$ is in K . By closure of \bar{K} under ultrapowers, $(\mathfrak{M}, w) \in K$ \square

Theorem 2.66. *Let τ be a modal similarity type, and K a class of pointed τ -models. Then the following are equivalent*

1. K is definable by means of a single modal formula
2. Both K and \bar{K} are closed under bisimulations and ultraproducts

Proof. Assume K, \bar{K} satisfy the stated conditions. Then both are closed under ultraproducts, hence by Theorem 2.65 there are set of modal formulas T_1, T_2 defining K and \bar{K} respectively. Observe their union is inconsistent in the sense that there is no pointed model (\mathfrak{M}, w) s.t. $(\mathfrak{M}, w) \models T_1 \cup T_2$. So by compactness there exists $\phi_1, \dots, \phi_n \in T_1$ and $\psi_1, \dots, \psi_m \in T_2$ s.t. for all pointed models (\mathfrak{M}, w)

$$\mathfrak{M}, w \models \phi_1 \wedge \dots \wedge \phi_n \rightarrow \neg\psi_1 \vee \dots \vee \neg\psi_m$$

By definition, for any $(\mathfrak{M}, w) \in K$ we have $\mathfrak{M}, w \models \phi_1 \wedge \dots \wedge \phi_n$. Conversely, if $\mathfrak{M}, w \models \phi_1 \wedge \dots \wedge \phi_n$, then $\mathfrak{M}, w \models \neg\psi_1 \vee \dots \vee \neg\psi_m$. Hence $\mathfrak{M}, w \not\models T_2$. Therefore $(\mathfrak{M}, w) \notin \bar{K}$, whence $(\mathfrak{M}, w) \in K$ \square

3 Frames

3.1 Frame Definability

Definition 3.1 (Validity). Let τ be a modal similarity type. A formula ϕ (of this similarity type) is **valid at a state w in a frame \mathfrak{F}** (notation: $\mathfrak{F}, w \models \phi$) if ϕ is true at w in every model (\mathfrak{F}, V) based on \mathfrak{F} ; ϕ is **valid on a frame \mathfrak{F}** (notation: $\mathfrak{F} \models \phi$) if it is valid at every state in \mathfrak{F} . A formula ϕ is **valid on a class of frames K** (notation: $K \models \phi$) if it is valid on every frame \mathfrak{F} in K . We denote the class of frames where ϕ is valid by Fr_ϕ

A set Γ of modal formulas (of type τ) is **valid on a frame \mathfrak{F}** if every formula in Γ is valid on \mathfrak{F} ; and Γ is **valid on a class K of frames** if Γ is valid on every member of K . We denote the class of frames where Γ is valid by Fr_Γ

Definition 3.2 (Definability). Let τ be a modal similarity type, ϕ a modal formula of this type, and K a class of τ -frames. We say that ϕ **defines** (or **characterizes**) K if for all frames \mathfrak{F} , \mathfrak{F} is in K iff $\mathfrak{F} \models \phi$. Similarly, if Γ is a set of modal formulas of this type, we say that Γ **defines** \mathfrak{F} is in K iff $\mathfrak{F} \models \Gamma$

A class of frames is **(modally) definable** if there is some set of modal formulas that defines it

Definition 3.3 (Relative Definability). Let τ be a modal similarity type, ϕ a modal formula of this type, and C a class of τ -frames. We say that ϕ **defines** (or **characterizes**) a class K of frames **within** C (or **relative to** C) if for all frames \mathfrak{F} in C we have that \mathfrak{F} is in K iff $\mathfrak{F} \models \phi$

Similarly, if Γ is a set of modal formulas of this type, we say that Γ **defines** a class K of frames **within** C if for all frames \mathfrak{F} in C we have that \mathfrak{F} is in K iff $\mathfrak{F} \models \Gamma$

Definition 3.4 (Frame Languages). For any modal similarity type τ , the **first-order frame language** of τ is the first-order language that has the identity symbol $=$ together with an $(n+1)$ -ary relation symbol R_Δ for each n -ary modal operator Δ in τ . We denote this language by \mathcal{L}_τ^1 . We often call it the **first-order correspondence language** (for τ)

Let Φ be any set of proposition letters. The **monadic second-order frame language** of τ over Φ is the monadic second-order language obtained by augmenting \mathcal{L}_τ^1 with a Φ -indexed collection of monadic predicate variables. (That is, this language has all the resources of \mathcal{L}_τ^1 , and in addition is capable of quantifying over subsets of frames). We denote this language by $\mathcal{L}_\tau^2(\Phi)$. We often simply call it the **second-order frame language** or the **second-order correspondence language** (for τ)

Definition 3.5 (Frame Correspondence). If a class of frames (property) can be defined by a modal formula ϕ and by a formula α from one of these frame languages, then we say that ϕ and α are each others (global) **frame correspondents**

For example the basic modal formula $p \rightarrow \Diamond p$ and the first-order sentence $\forall x Rxx$ are correspondents

Example 3.1. Read $\Diamond\phi$ as 'it is **possibly** the case that ϕ ' and $\Box\phi$ as '**necessarily** ϕ '.

$$(T) \quad p \rightarrow \Box p$$

$$(4) \quad \Diamond\Diamond p \rightarrow \Diamond p$$

$$(5) \Diamond p \rightarrow \Box \Diamond p$$

Our first claim is that for any frame $\mathfrak{F} = (W, R)$, the axiom T corresponds to **reflexivity** of the relation R :

$$\mathfrak{F} \models T \quad \text{iff} \quad \mathfrak{F} \models \forall x Rxx$$

Suppose that R is **not** reflexive. There exists a state w which is not accessible from itself. Now the valuation V has to satisfy two conditions

1. $w \in V(p)$
2. $\{x \in W \mid Rwx\} \cap V(p) = \emptyset$

Consider the **minimal** valuation V satisfying condition (1), that is, take

$$V(p) = \{w\}$$

Now let v be an R -successor of w . As Rww does not hold in \mathfrak{F} , v must be distinct from w , so $v \notin V(p)$. As v was arbitrary, $w \notin V(p)$

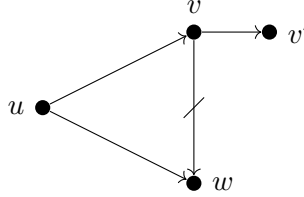
Likewise, one can prove that for any frame $\mathfrak{F} = (W, R)$

$$\mathfrak{F} \models 4 \quad \text{iff} \quad R \text{ is transitive}$$

$$\mathfrak{F} \models 5 \quad \text{iff} \quad R \text{ is euclidean}$$

where a relation is **euclidean** if it satisfies $\forall xyz((Rxy \wedge Rxz) \rightarrow Ryz)$.

Assume \mathfrak{F} is a non-euclidean frame; then there must be states u, v, w



s.t. Ruv, Ruw but not Rvw :

We will try to falsify 5 in u ; for this purpose we have to find a valuation V s.t. $(\mathfrak{F}, V), u \models \Diamond p$ and $(\mathfrak{F}, V), u \not\models \Box \Diamond p$. In other words, we have to make p **true** at some R -successor x of u , and **false** at all R -successors of some R -successor y of u . The constraints on V are

1. $w \in V(p)$
2. $\{z \mid Rvz\} \cap V(p) = \emptyset$

Let's take a **maximal** V satisfying condition (2), that is, define

$$V(p) = \{z \in W \mid \text{it is not the case that } Rvz\}$$

Now $v \not\models \Diamond p$, so $u \not\models \Box \Diamond p$. On the other hand, we have $w \models p$, so $u \models \Diamond p$

Example 3.2. Suppose that we are working with the basic temporal language and that we are interested in **dense** bidirectional frames. This property can be defined using a first-order sentence (namely $\forall xy(x < y \rightarrow \exists z(x < z \wedge z < y))$) but can the basic temporal language define it too?

The following simple formula suffices: $Fp \rightarrow FFp$. Let $\mathfrak{T} = (T, <)$ be a frame s.t. $\mathfrak{T} \models Fp \rightarrow FFp$. Suppose that a point $t \in T$ has a $<$ -successor t' . Consider the following **minimal** valuation V_m guaranteeing that $(\mathfrak{T}, V_m), t \models Fp$

$$V_m(p) = \{t'\}$$

Hence $t \models FFp$. This means there is a point s s.t. $t < s$ and $s \models Fp$. But as t' is the **only** states where p holds, this implies that $s < t'$

Example 3.3. Suppose we are working with a similarity type with three binary operators $\triangle_1, \triangle_2, \triangle_3$, and that we are interested in the class of frames in which the three ternary accessibility relations (denoted by R_1, R_2, R_3 respectively, if $R_1 stu$ and $s \models p, t \models q, s \models r$, then $s \models q\triangle_1 r$).

We want

$$R_1 stu \quad \text{iff} \quad R_2 tus \quad \text{iff} \quad R_3 ust$$

to hold for all s, t, u in such frames. Can we define this class of frames?

We can. We will show that for all frames $\mathcal{F} = (W, R_1, R_2, R_3)$ we have

$$\mathcal{F} \models p \wedge (q\triangle_1 r) \rightarrow (q \wedge r\triangle_2 p)\triangle_1 r \quad \text{iff} \quad \mathcal{F} \models \forall xyz(R_1 xyz \rightarrow R_2 yzx)$$

Suppose that the modal formula $p \wedge (q\triangle_1 r) \rightarrow (q \wedge r\triangle_2 p)\triangle_1 r$ is valid in \mathcal{F} , and consider states s, t, u with $R_1 stu$. Consider a valuation V with $V(p) = \{s\}, V(q) = \{t\}, V(r) = \{u\}$. Then $(\mathcal{F}, V), s \models p \wedge q\triangle_1 r$, so $s \models (q \wedge r\triangle_2 p)\triangle_1 r$. Hence there must be states t', u' with $R_1 st'u', t' \models q \wedge r\triangle_2 p$ and $u' \models r$.

Exercise 3.1.1. Consider a language with two diamonds $\langle 1 \rangle$ and $\langle 2 \rangle$. Show that $p \rightarrow [2]\langle 1 \rangle p$ is valid on precisely those frame for the language satisfy the condition $\forall xy(R_2 xy \rightarrow R_1 yx)$. What sort of frames does $p \rightarrow [1]\langle 1 \rangle p$ define?

Proof. Define $V(p) = \{w\}$ to prove left to right.

$$R_1 xy \rightarrow R_1 yx$$

□

Exercise 3.1.2. Consider a language with three diamonds $\langle 1 \rangle, \langle 2 \rangle, \langle 3 \rangle$. Show that the modal formula $\langle 3 \rangle p \leftrightarrow \langle 1 \rangle \langle 2 \rangle p$ is valid on a frame for this language iff the frame satisfies the condition $\forall xy(R_3 xy \leftrightarrow \exists z(R_1 xz \wedge R_2 zy))$

3.2 Frame Definability and Second-Order Logic

Example 3.4. Consider the Löb formula $\Box(\Box p \rightarrow p) \rightarrow \Box p$, which we will call it L for brevity. We show that L defines the class of frames (W, R) s.t. R is transitive and R 's converse is well-founded

We will then show that this is a class of frames that first-order frame languages **cannot** define; that is, we will show that this class is not elementary

Assume that $\mathfrak{F} = (W, R)$ is a frame with a transitive and conversely well-founded relation, and then suppose that L is not valid in \mathfrak{F} . This means that there is a valuation V and a state w s.t. $(\mathfrak{F}, V), w \not\models \Box(\Box p \rightarrow p) \rightarrow \Box p$. In other words, $w \models \Box(\Box p \rightarrow p)$ but $w \not\models \Box p$. Then w must have a successor w_1 s.t. $w_1 \not\models p$, and as $w_1 \models \Box p \rightarrow p$, we have $w_1 \not\models \Box p$. This in turn implies that w_1 have a successor w_2 where p is false; note that by the transitivity of R , w_2 is also a successor of w . Again, w_2 must have a p -falsifying successor w_3 . Hence we find an infinite path $wRw_1Rw_2R\dots$ contradicting the converse well-foundedness of R

For the other direction, assume that either R is not transitive or its converse is not well-founded; in both cases we have to find a valuation V and a state w s.t. $(\mathfrak{F}, V), w \not\models L$. Assume that R is transitive, but not conversely well-founded. In other words, suppose we have a transitive frame containing an infinite sequence $w_0Rw_1Rw_2R\dots$. Define

$$V(p) = W \setminus \{x \in W \mid \text{there is an infinite path starting from } x\}$$

$\Box p \rightarrow p$ is true **everywhere** in the model, whence certainly, $(\mathfrak{F}, V), w_0 \models \Box(\Box p \rightarrow p)$. The claim then follows from the fact that $(\mathfrak{F}, V), w_0 \not\models \Box p$

Assume that R is not transitive and its converse is well-founded. Since R is not transitive, there is Rw_1w_2 and Rw_2w_3 but $\neg Rw_1w_3$. Let $V(p) = \{w_2, w_3\}$. Then $(\mathfrak{F}, V), w_1 \models \Diamond p$ and $(\mathfrak{F}, V), w_1 \models \neg \Diamond(p \wedge \neg \Diamond p)$

Finally, to show that the class of frames defined by L is not elementary, an easy compactness argument suffices. Suppose for the sake of a contradiction that there is a first-order formula equivalent to L ; call this formula λ . As λ is equivalent to L , and model making λ true must be transitive. Let $\sigma_n(x_0, \dots, x_n)$ be the first-order formula stating that there is an R -path of length n through x_0, \dots, x_n :

$$\sigma_n(x_0, \dots, x_n) = \bigwedge_{0 \leq i < n} Rx_i x_{i+1}$$

Every **finite** subset of

$$\Sigma = \{\lambda\} \cup \{\forall xyz((Rxy \wedge Ryz) \rightarrow Rxz)\} \cup \{\sigma_n \mid n \in \omega\}$$

is satisfiable in a finite linear order, and hence in the class of transitive, conversely well-founded frames. Thus Σ must have a model. But it is clear that Σ is **not** satisfiable in any conversely well-founded frame.

Example 3.5. PDL can be interpreted on any transition system of the form $\mathcal{F} = (W, R_\pi)_{\pi \in \Pi}$. Let's call such a frame ***-proper** if the transition relation R_{π^*} of each program π^* is the reflexive and transitive closure of the transitive relation R_π of π .

Consider the following set of formulas

$$\Delta = \{[\pi]^*(p \rightarrow [\pi]p) \rightarrow (p \rightarrow [\pi^*]p), \langle \pi^* \rangle p \leftrightarrow (p \vee \langle \pi \rangle \langle \pi^* \rangle p) \mid \pi \in \Pi\}$$

called **Segerberg's axiom**, or the **induction axiom**. We claim that for any PDL-frame \mathfrak{F}

$$\mathfrak{F} \models \Delta \quad \text{iff} \quad \mathfrak{F} \text{ is } \ast\text{-proper}$$

A consequences is that PDL is strong enough to define the class of regular frames. The constraints on the relations interpreting \cup and $;$ are simple first-order conditions, and

$$\Gamma = \{\langle \pi_1; \pi_2 \rangle p \leftrightarrow \langle \pi_1 \rangle \langle \pi_2 \rangle p, \langle \pi_1 \cup \pi_2 \rangle p \leftrightarrow \langle \pi_1 \rangle p \vee \langle \pi_2 \rangle p \mid \pi \in \Pi\}$$

pins down what is required. So $\Delta \cup \Gamma$ defines the regular frames

Example 3.6. We will show that the McKinsey formula (M) $\Box \Diamond p \rightarrow \Diamond \Box p$ does not correspond to a first-order condition by show that it violates the Löwenheim-Skolem Theorem

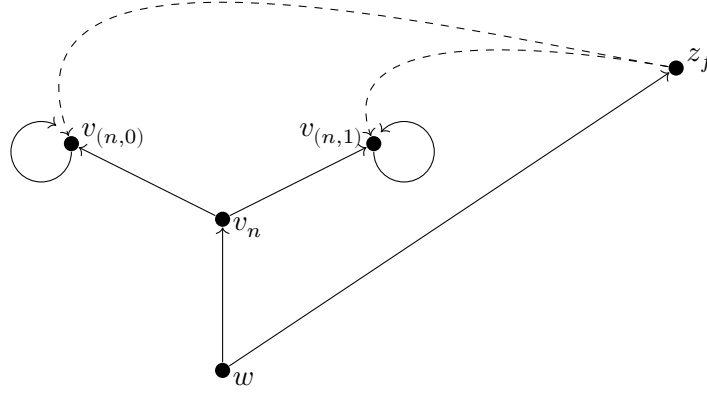
Consider the frame $\mathfrak{F} = (W, R)$, where

$$W = \{w\} \cup \{v_n, v_{(n,i)} \mid n \in \mathbb{N}, i \in \{0, 1\}\} \cup \{z_f \mid f : \mathbb{N} \rightarrow \{0, 1\}\}$$

and

$$R = \{(w, v_n), (v_n, v_{(n,i)}), (v_{(n,i)}, v_{(n,i)}) \mid n \in \mathbb{N}, i \in \{0, 1\}\} \cup \{(w, z_f), (z_f, v_{(n,f(n))}) \mid n \in \mathbb{N}, f : \mathbb{N} \rightarrow \{0, 1\}\}$$

In a picture



Note that $|W| = \aleph_1$

Our first observation is that $\mathfrak{F} \Vdash \Box \Diamond p \rightarrow \Diamond \Box p$.

If $\mathfrak{F}, v_n \Vdash \Box \Diamond p$, then both two $\mathfrak{F}, v_{(n,i)} \Vdash \Diamond p$, which means $\{v_{(n,0)}, v_{(n,1)}\} \subseteq V(p)$. Hence $\mathfrak{F}, v_{(n,i)} \Vdash \Box p$ and $\mathfrak{F}, v_n \Vdash \Box \Diamond p \rightarrow \Diamond \Box p$.

If $\mathfrak{F}, z_f \Vdash \Box \Diamond p$, then for each $n \in \mathbb{N}$, there is a $i = 0$ or 1 s.t. $\mathfrak{F}, v_{(n,i)} \Vdash \Diamond p$. Then $\mathfrak{F}, v_{(n,i)} \Vdash p$

If $\mathfrak{F}, w \Vdash \Box \Diamond p$, then either $v_{(n,0)} \in V(p)$ or $v_{(n,1)} \in V(p)$. Choose $f : \mathbb{N} \rightarrow \{0, 1\}$ s.t. $\mathfrak{F}, v_{(n,f(n))} \Vdash p$ for each $n \in \mathbb{N}$. Then clearly $\mathfrak{F}, z_f \Vdash \Box p$, and so $\mathfrak{F}, w \Vdash \Diamond \Box p$

In order to show that $\Box \Diamond p \rightarrow \Diamond \Box p$ does not define a first-order frame condition, let us view the frame \mathfrak{F} as a first-order model with domain W . By the downward Löwenheim-Skolem Theorem, there must be a countable elementary submodel \mathfrak{F}' of \mathfrak{F} whose domain W' contains w and each $v_n, v_{(n,0)}, v_{(n,1)}$. As W is uncountable and W' countable, there must be a mapping $f : \mathbb{N} \rightarrow \{0, 1\}$ s.t. $z_f \notin W'$. Now if the McKinsey formula was equivalent to a first-order formula it would be valid on \mathfrak{F}' . But we will show that the McKinsey formula is **not** valid on \mathfrak{F}' , hence it cannot be equivalent to a first-order formula

Let V' be a valuation on \mathfrak{F}' s.t. $V'(p) = \{v_{(n,f(n))} \mid n \in \mathbb{N}\}$, here f is a mapping s.t. $z_f \notin W'$.

First, $(\mathfrak{F}', V'), w \not\Vdash \Diamond \Box p$.

Then we need to show that $(\mathfrak{F}', V'), w \Vdash \Box \Diamond p$. Consider arbitrary element z_g of W' . Call two states z_h and z_k of \mathfrak{F} **complementary** if for all n , $h(n) = 1 - k(n)$. Now suppose z_g is complement to z_f ; since complementary states are unique, the fact that \mathfrak{F}' is an elementary submodel of \mathfrak{F} would imply that z_f exists in \mathfrak{F}' ($\mathfrak{F}' \models \forall f \exists g (f, g \text{ are complement})$). Hence we may

conclude that z_g is **not** complementary to z_f . Hence there exists some $n \in \mathbb{N}$ s.t. $g(n) = f(n)$. Therefore $(\mathfrak{F}', V'), z_g \Vdash \Diamond p$. Then $(\mathfrak{F}', V'), w \Vdash \Box \Diamond p$

View the predicate symbol P that corresponds to the proposition letter p as a monadic second-order variable that we can quantify over

Proposition 3.6. *Let τ be a modal similarity type, and ϕ a τ -formula. Then for any τ -frames and any state w in \mathfrak{F}*

$$\begin{aligned} \mathfrak{F}, w \Vdash \phi & \quad \text{iff} \quad \mathfrak{F} \models \forall P_1 \dots \forall P_n ST_x(\phi)[w] \\ \mathfrak{F} \Vdash \phi & \quad \text{iff} \quad \mathfrak{F} \models \forall P_1 \dots \forall P_n \forall x ST_x(\phi) \end{aligned}$$

Here, the second-order quantifier bind second-order variables P_i corresponding to the proposition letters p_i occurring in ϕ

Proof. Let $\mathfrak{M} = (\mathfrak{F}, V)$ be any model based on \mathfrak{F} , and let w be any state in \mathfrak{F} . Then we have that

$$(\mathfrak{F}, V), w \Vdash \phi \quad \text{iff} \quad \mathfrak{F} \models ST_x(\phi)[w, P_1, \dots, P_n]$$

where the notation $[w, P_1, \dots, P_n]$ means ‘assign w to the free first-order variable x in $ST_x(\phi)$, and $V(p_1), \dots, V(p_n)$ to the free monadic second-order variables.’ \square

Refer to $\forall P_1 \dots \forall P_n \forall x ST_x(\phi)$ as the **second-order translation** of ϕ

A general frame can be viewed as a **generalized model** for (monadic) second-order logic. A generalized model for second-order logic is a model in which the second-order quantifiers are viewed as ranging not over **all** subsets, but only over a pre-selected sub-collection of subsets. This means the following equivalence holds

$$(\mathfrak{F}, A) \Vdash \phi \quad \text{iff} \quad (\mathfrak{F}, A) \models \forall P_1 \dots \forall P_n \forall x ST_x(\phi)$$

Exercise 3.2.1. 1. Consider a modal language with two diamonds $\langle 1 \rangle$ and $\langle 2 \rangle$. Prove that the class of frames in which R_1 is the reflexive transitive closure of R_2 is defined by the conjunction of the formulas $\langle 1 \rangle p \rightarrow (p \vee \langle 1 \rangle (\neg p \wedge \langle 2 \rangle p))$ and $\langle 1 \rangle p \leftrightarrow (p \vee \langle 2 \rangle \langle 1 \rangle p)$

Proof. 1. **Right to left.** $p \vee \langle 2 \rangle \langle 1 \rangle p \rightarrow \langle 1 \rangle p$ means $p \rightarrow \langle 1 \rangle p$ and $\langle 2 \rangle \langle 1 \rangle p \rightarrow \langle 1 \rangle p$. Hence the frame must be reflexive.

By reflexivity and $\langle 2 \rangle \langle 1 \rangle p \rightarrow \langle 1 \rangle p$, we can show that $R_2 xy \rightarrow R_1 xy$. Consider

$$x \Vdash \neg p \xrightarrow{R_2} y \Vdash p \xleftarrow{\quad} R_1$$

and $V(p) = \{y\}$. Since $x \Vdash \langle 2 \rangle \langle 1 \rangle p$, we have $x \Vdash \langle 1 \rangle p$. Hence we must have $R_1 xy$

Left to right. If R_1 is the transitive closure of R_2 . $p \rightarrow \langle 1 \rangle p$ and $\langle 2 \rangle \langle 1 \rangle p \rightarrow \langle 1 \rangle p$ is obvious. □

Exercise 3.2.2. Show that Grzegorzczuk's formula, $\Box(\Box(p \rightarrow \Box p) \rightarrow p) \rightarrow p$, characterizes the class of frames $\mathfrak{F} = (W, R)$ satisfying

1. R is reflexive
2. R is transitive
3. there is no infinite paths $x_0 R x_1 R x_2 R \dots$ s.t. for all $i, x_i \neq x_{i+1}$

Proof. Suppose $\mathfrak{F} \Vdash \Box(\Box(p \rightarrow \Box p) \rightarrow p) \rightarrow p$.

1. If R is not reflexive, we need to find a valuation s.t. $(\mathfrak{F}, V), w \Vdash \Box(\Box(p \rightarrow \Box p) \rightarrow p)$ but $(\mathfrak{F}, V), w \nVdash p$. This is done by $V(p) = W - \{w\}$
2. If R is reflexive but not transitive, then there is $w R u R v$ but not $w R v$. Let V be an evaluation on \mathfrak{F} s.t. $V(p) = W - \{w, v\}$. Then $(\mathfrak{F}, V), u \Vdash p \wedge \Diamond \neg p$ and hence $(\mathfrak{F}, V), w \Vdash \Diamond(p \wedge \Diamond \neg p)$. This is equivalent to $(\mathfrak{F}, V), w \nVdash \Box(p \rightarrow \Box p)$, from which it follows that $(\mathfrak{F}, V), w \Vdash \Box(p \rightarrow \Box p) \rightarrow p$. Thus for all $x \in w \uparrow$, $(\mathfrak{F}, V), x \Vdash \Box(p \rightarrow \Box p) \rightarrow p$ and hence $(\mathfrak{F}, V), w \Vdash \Box(\Box(p \rightarrow \Box p) \rightarrow p)$ but $(\mathfrak{F}, V), w \nVdash p$
3. If R is reflexive and transitive, but there is some infinite path $x_0 R x_1 R x_2 R \dots$ s.t. for all $i, x_i \neq x_{i+1}$, then let V be an evaluation on \mathfrak{F} s.t. $V(p) = W - \{x_{2k} \mid k \in \mathbb{N}\}$. For all x_{2k} with some $k \in \mathbb{N}$, $(\mathfrak{F}, V), x_{2k+2} \Vdash \neg p$ and $(\mathfrak{F}, V), x_{2k+1} \Vdash p$. It follows that $(\mathfrak{F}, V), x_{2k+1} \Vdash p \wedge \Diamond \neg p$ and $(\mathfrak{F}, V), x_{2k} \Vdash \Diamond(p \wedge \Diamond \neg p)$. Hence $(\mathfrak{F}, V), x_{2k} \Vdash \Box(p \rightarrow \Box p) \rightarrow p$. For all $u \in W - \{x_{2k} \mid k \in \mathbb{N}\}$, $(\mathfrak{F}, V), u \Vdash \Box(p \rightarrow \Box p) \rightarrow p$

Now to prove $\mathfrak{F} \Vdash \Box(\Box(p \rightarrow \Box p) \rightarrow p) \rightarrow p$. Suppose there is an evaluation V and a point w s.t. $(\mathfrak{F}, V), w \nVdash \Box(\Box(p \rightarrow \Box p) \rightarrow p) \rightarrow p$. It follows $(\mathfrak{F}, V), w \Vdash \Box(\Box(p \rightarrow \Box p) \rightarrow p)$ and $(\mathfrak{F}, V), w \nVdash p$. Then for all $u \in w \uparrow$, $(\mathfrak{F}, V), u \Vdash \Box(p \rightarrow \Box p) \rightarrow p$. Since R is reflexive and transitive, $(\mathfrak{F}, V), w \Vdash \Box(p \rightarrow \Box p) \rightarrow p$. It follows that $(\mathfrak{F}, V), w \Vdash \Diamond(p \wedge \neg \Box p)$. Hence there is a point $v \in w \uparrow$ s.t. $(\mathfrak{F}, V), v \Vdash p \wedge \neg \Box p$. Since $(\mathfrak{F}, V), w \Vdash \neg p$, $v \neq w$.

And there is some point $v' \in v \uparrow$, $(\mathfrak{F}, V), v' \Vdash \neg p$, this also shows that $v' \neq v$. By the transitivity, $v' \in w \uparrow$, $(\mathfrak{F}, V), v' \Vdash \Box(p \rightarrow \Box p) \rightarrow p$. Similarly we can find a distinct point $v'' \in v' \uparrow$ s.t. $(\mathfrak{F}, V), v'' \Vdash p \wedge p \wedge \neg \Box p$. By repeating this line of reasoning, we can find an infinite path in which each point is different from its successor. This contradicts 3. \square

Exercise 3.2.3. Consider the basic temporal language. Recall that a frame $\mathfrak{F} = (W, R_F, R_P)$ for this language is called **bidirectional** if R_P is the converse of R_F

1. Prove that among the finite bidirectional frames, the formula $G(Gp \rightarrow p) \rightarrow Gp$ together with its converse, $H(Hp \rightarrow p) \rightarrow Hp$ defines the transitive and irreflexive frames

Proof. 1. Let $\mathfrak{F} = (W, R)$ be an finite frame and suppose that \mathfrak{F} is transitive and irreflexive. Suppose $(\mathfrak{F}, V), w \Vdash G(Gp \rightarrow p)$ and $(\mathfrak{F}, V), w \Vdash \neg Gp$. We can find an infinite chain

If \mathfrak{F} is not transitive, then there are points w, u, v s.t. Rwu and Ruv but not in the case that Rwv . Let V be an evaluation on \mathfrak{F} s.t. $V(p) = W - \{u, v\}$. Then $(\mathfrak{F}, V), v \Vdash \neg p$ and $(\mathfrak{F}, V), u \Vdash \neg Gp$ and hence $(\mathfrak{F}, V), u \Vdash Gp \rightarrow p$. Since $v \notin w \uparrow$, we have $(\mathfrak{F}, V), w \Vdash G(Gp \rightarrow p)$

\square

3.3 Definable and Undefinable Properties

Definition 3.7. A bounded morphism from a τ -frame $\mathfrak{F} = (W, R_\Delta)_{\Delta \in \tau}$ to a τ -frame $\mathfrak{F}' = (W', R'_\Delta)_{\Delta \in \tau}$ is a function from W to W' satisfying the following two conditions

1. for all $\Delta \in \tau$, $R_\Delta w v_1 \dots v_n$ implice $R'_\Delta f(w) f(v_1) \dots f(v_n)$
2. if $R'_\Delta f(w) v'_1 \dots v'_n$ then there exist $v_1 \dots v_n$ s.t. $R_\Delta w v_1 \dots v_n$ and $f(v_i) = v'_i$ for $1 \leq i \leq n$

Theorem 3.8. Let τ be a modal similarity type, and ϕ a τ -formula

1. Let $\{\mathfrak{F}_i \mid i \in I\}$ be a family of frames. Then $\biguplus \mathfrak{F}_i \Vdash \phi$ if $\mathfrak{F}_i \Vdash \phi$ for every i in I
2. Assume that $\mathfrak{F}' \rightarrow \mathfrak{F}$. Then $\mathfrak{F}' \Vdash \phi$ if $\mathfrak{F} \Vdash \phi$
3. Assume that $\mathfrak{F} \rightarrow \mathfrak{F}'$. Then $\mathfrak{F} \Vdash \phi$ if $\mathfrak{F}' \Vdash \phi$

Proof. 1. Assume that $\mathfrak{F}_i \Vdash \phi$ for every i in I but $\bigsqcup_i \mathfrak{F}_i \nVdash \phi$. Then there is a valuation V and a state w s.t. $(\bigsqcup_i \mathfrak{F}_i, V), w \Vdash \neg\phi$. If w is a state in \mathfrak{F}_i , then define a valuation V' on \mathfrak{F}_i

$$V'(p_i) = \{x \in W_i \mid x \in V(p_i)\}$$

Then $(\mathfrak{F}_i, V'), w \nVdash \phi$

2. Suppose there is a valuation and a state in \mathfrak{F}' s.t. $(\mathfrak{F}', V), w \nVdash \phi$. Then $(\mathfrak{F}, V), w \nVdash \phi$
3. Assume that f is a surjective bounded morphism from \mathfrak{F} onto \mathfrak{F}' and that $\mathfrak{F} \Vdash \phi$. We had to show that $\mathfrak{F}' \Vdash \phi$. Suppose that ϕ is not valid in \mathfrak{F}' . Then there must be a valuation V' and a state w' s.t. $(\mathfrak{F}', V'), w' \nVdash \phi$. Define the following valuation V on \mathfrak{F} :

$$V(p_i) = \{x \in W \mid f(x) \in V'(p_i)\}$$

It follows that $(\mathfrak{F}, V), w \nVdash \phi$

□

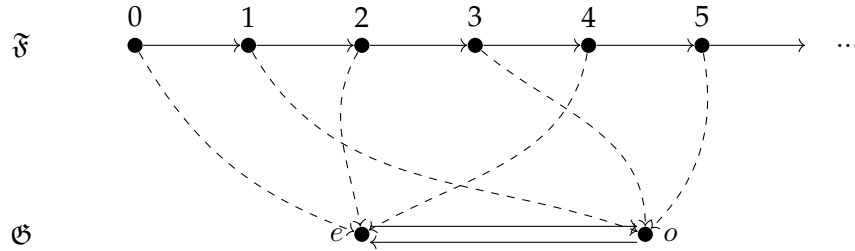
Example 3.7. The class of finite frames is not modally definable. For suppose there was a set of formulas Δ characterizing the finite frames. Then Δ would be valid in every one-point frame $\mathfrak{F}_i = \{(\{w_i\}, \{(w_i, w_i)\})\} (i < \omega)$. By Theorem 3.8 this would imply that Δ was also valid in the disjoint union $\bigsqcup_i \mathfrak{F}_i$

The class of frames having a reflexive point ($\exists x Rxx$) does not have a modal characterization either. For suppose that the set Δ characterized this class. Consider the following frame

$$w \curvearrowright \quad v \longleftrightarrow u$$

As w is reflexive, $\mathfrak{F} \Vdash \Delta$. Now consider the generated subframe \mathfrak{F}_v of \mathfrak{F} .

Consider the following two frames: $\mathfrak{F} = (\omega, S)$, the natural numbers with the successor relation and $\mathfrak{G} = (\{e, o\}, \{(e, o), (o, e)\})$ as



No property P is modally definable if \mathfrak{F} has P and \mathfrak{G} lacks. This shows that there is no set of formulas characterizing the asymmetric frames $\forall(Rxy \rightarrow \neg Ryx)$

Corollary 3.9. *Let τ be a modal similarity type, \mathfrak{F} a τ -frame, and ϕ a τ -formula. Then $\mathfrak{F} \models \phi$ if $\text{ue}\mathfrak{F} \models \phi$*

Proof. Assume that ϕ is not valid in \mathfrak{F} . That is, there is a valuation V and a state w s.t. $(\mathfrak{F}, V), w \models \neg\phi$. By Proposition 2.52 $\neg\phi$ is true at u_w in the ultrafilter extension of \mathfrak{M} . \square

We claim that every state has a reflexive successor $\forall x\exists y(Rxy \wedge Ryy)$ is not modally definable, even though it is preserved under taking disjoint unions, generated subframes and bounded morphic images. Consider the frame in Example 2.4.

Theorem 3.10. *Let τ be a modal similarity type, and \mathfrak{F} a τ -frame. Then \mathfrak{F} has an ultrapower $\prod_U \mathfrak{F}$ s.t. $\prod_U \mathfrak{F} \twoheadrightarrow \text{ue}\mathfrak{F}$. In a diagram*

$$\begin{array}{ccc} \prod_U \mathfrak{F} & & \\ \downarrow & \searrow & \\ \mathfrak{F} & & \text{ue}\mathfrak{F} \end{array}$$

Corollary 3.11. *Let τ be a modal similarity type, and ϕ a τ -formula. If ϕ defines a first-order property of frames, then frame validity of ϕ is preserved under taking ultrafilter extensions*

Proof. Let ϕ be a modal formula which defines a first-order property of frames, and let \mathfrak{F} be a frame s.t. $\mathfrak{F} \models \phi$. By the previous theorem, there is an ultrapower $\prod_U \mathfrak{F}$ of \mathfrak{F} s.t. $\prod_U \mathfrak{F} \twoheadrightarrow \text{ue}\mathfrak{F}$. As first-order properties are preserved under ultrapowers, $\prod_U \mathfrak{F} \models \phi$. But then $\text{ue}\mathfrak{F} \models \phi$. \square

Theorem 3.12 (Goldblatt-Thomason Theorem). *Let τ be a modal similarity type. A first-order definable class K of τ -frames is modally definable iff it is closed under taking bounded morphic images, generated subframes, disjoint unions and reflects ultrafilter extensions*

Exercise 3.3.1. Consider the basic modal language. Show that the following properties of frames are not modally definable

1. antisymmetry $(\forall xy)(Rxy \wedge Ryx \rightarrow x = y)$
2. $|W| > 23$

3. $|W| < 23$
4. acyclicity
5. every state has at most one predecessor
6. every state has at least two successors

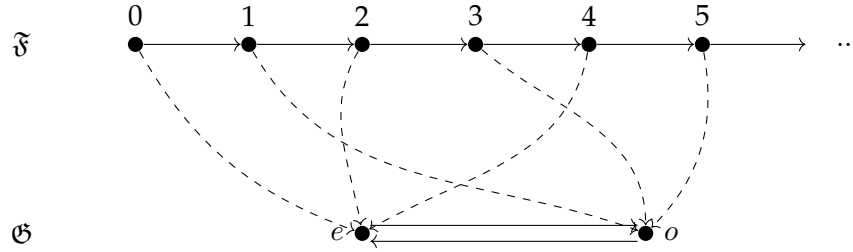
Proof. 1. Consider the frame in the example of bounded morphism

2. bounded morphism

3. Disjoint union

4. bounded morphism. Consider the following two frames: $\mathfrak{F} = (\omega, S)$, the natural numbers with the successor relation $S = \{(n, m) \mid m = n + 1\}$ and

$\mathfrak{G} = (\{e, o\}, \{(e, o), (o, e)\})$ as



5. bounded morphism. Consider two frames

$\mathfrak{M} = (\{a, b, c, d\}, \{(a, c), (b, d)\})$ and $\mathfrak{N} = (\{a', b', c'\}, \{(a', c'), (b', c')\})$ with surjective bounded morphism $f(a) = a', f(b) = b', f(c) = f(d) = c'$. In \mathfrak{M} every state has at most one predecessor, but not in \mathfrak{N}

6. bounded morphism. Consider two frames

$\mathfrak{M} = (\{a, b, c\}, \{(a, b), (a, c), (b, a), (b, c), (c, a), (c, b)\})$ and $\mathfrak{N} = (\{e\}, \{(e, e)\})$ with bounded morphism $f(a) = f(b) = f(c) = e$. In \mathfrak{M} , every state has two successors but in \mathfrak{N} there is only one

□

3.4 Finite Frames

3.4.1 Finite transitive frames

Let $\mathfrak{F} = (W, R)$ be a point-generated finite transitive frame for the basic modal similarity type, and let w be a root of \mathfrak{F} . The Jankov-Fine formula $\phi_{\mathfrak{F}, w}$ is a description of \mathfrak{F} that has the following property: it is satisfiable on a frame iff \mathfrak{F} is a bounded morphic image of a generated subframe of \mathfrak{G}

We build Jankov-Fine formulas as follows. Enumerate the states of \mathfrak{F} as w_0, \dots, w_n , where $w = w_0$. Associate each state w_i with a distinct proposition letter p_i . Let $\phi_{\mathfrak{F}, w}$ be the conjunction of the following formulas:

1. p_0
2. $\Box(p_0 \vee \dots \vee p_n)$
3. $(p_i \rightarrow \neg p_j) \wedge \Box(p_i \rightarrow \neg p_j)$ for each $i \neq j \leq n$
4. $(p_i \rightarrow \Diamond p_j) \wedge \Box(p_i \rightarrow \Diamond p_j)$ for each i, j with $Rw_i w_j$
5. $(p_i \rightarrow \neg \Diamond p_j) \wedge \Box(p_i \rightarrow \neg \Diamond p_j)$ for each i, j with $\neg Rw_i w_j$

Note that as R is transitive, each node in \mathfrak{F} is accessible in one step from w . It followss that when formulas of the form $\psi \wedge \Box\psi$ are satisfied at w , ψ is true throughout \mathfrak{F} .

Lemma 3.13. *Let \mathfrak{F} be a transitive, finite, point-generated frame, let w be a root of \mathfrak{F} , and let $\phi_{\mathfrak{F}, w}$ be the Jankov-Fine formula for \mathfrak{F} and w . Then for any transitive frame \mathfrak{G} we have the following equivalence: there is a valuation V and a node v s.t. $(\mathfrak{G}, V), v \models \phi_{\mathfrak{F}, w}$ iff there exists a bounded morphism from \mathfrak{G}_v onto \mathfrak{F}*

Proof. Left to right. Suppose there is a valuation V and a node v s.t. $(\mathfrak{G}, V), v \models \phi_{\mathfrak{F}, w}$. 1 shows that $v \in V(p_0)$. 2 and 3 shows that each of states satisfies exactly one of proposition letter p_i . Hence we build a function $f : \mathfrak{G}_v \rightarrow \mathfrak{F}$ as $f(v_i) = w_j$ if $v_i \in V(p_j)$ for each $v_i \in \mathfrak{G}$. 2 and 3 show that this function is well-defined.

Suppose $R_{\mathfrak{G}} v_i v_j$ and $\neg R_{\mathfrak{F}} w_i w_j$, then $(\mathfrak{G}, V), g \models p_i \rightarrow \neg \Diamond p_j$ for all $g \in \mathfrak{G}$. However, $(\mathfrak{G}, V), v_i \models p_i$ and $(\mathfrak{G}, V), v_j \models p_j$

If $R_{\mathfrak{F}} f(v_i) w_j$, then $(\mathfrak{G}, V), v_i \models p_{f(v_i)} \rightarrow \Diamond p_j$. Hence there is a successor v_j of v_i s.t. $f(v_j) = w_j$. Therefore we find $R_{\mathfrak{G}} v_i v_j$ and f is a bounded morphism

Right to left. Suppose there is a bounded morphism f from \mathfrak{G}_v onto \mathfrak{F} . Then clearly $f(v) = w_0$. Let $V(p_{f(v_i)}) = \{v_j \mid f(v_j) = f(v_i)\}$ for all $v_i \in \mathfrak{G}$. Then $(\mathfrak{G}, V), v \models p_0 \wedge \Box(p_0 \vee \dots \vee p_n) \wedge (p_i \rightarrow \neg p_j) \wedge \Box(p_i \rightarrow \neg p_j)$. If $R_{\mathfrak{F}} w_i w_j$

and there is a $v_i \in \mathfrak{G}$ s.t. $f(v_i) = w_i$, then there is a $f(v_j) = w_j$ s.t. $R_{\mathfrak{G}} v_i v_j$. Hence $(\mathfrak{F}, V), v \Vdash (p_i \rightarrow \Diamond p_j) \wedge \Box(p_i \rightarrow \Diamond p_j)$ for each i, j with $Rw_i w_j$. 5 is the same \square

Theorem 3.14. *Recall that τ_0 denotes the basic modal similarity type. Let K be a class of finite, transitive τ_0 -frames. Then K is definable within the class of finite, transitive τ_0 -frames iff it is closed under taking (finite) disjoint unions, generated subframes and bounded morphic images*

Proof. Assume that K is a class of finite, transitive τ_0 -frames and satisfies the stated closure conditions. Let Λ_K be the logic of K ; that is, $\Lambda_K = \{\phi \mid \mathfrak{F} \Vdash \phi, \text{ for all } \mathfrak{F} \in K\}$. We will show that Λ_K defines K . We need to show that if $\mathfrak{F} \Vdash \Lambda_K$, where \mathfrak{F} is finite and transitive, then $\mathfrak{F} \in K$.

First suppose \mathfrak{F} is point-generated with root w . Consider the Jankov-Fine formula $\phi_{\mathfrak{F}, w}$ for \mathfrak{F} and w . Clearly $\phi_{\mathfrak{F}, w}$ is satisfiable in \mathfrak{F} at w , so $\neg\phi_{\mathfrak{F}, w} \notin \Lambda_K$. Hence there is some $\mathfrak{G} \in K$ s.t. $\mathfrak{G} \not\Vdash \neg\phi_{\mathfrak{F}, w}$; hence there is some valuation V and state v s.t. $(\mathfrak{G}, V), v \Vdash \phi_{\mathfrak{F}, w}$. Thus by the previous lemma, \mathfrak{F} is a bounded morphic image of the point-generated subframe \mathfrak{G}_v of \mathfrak{G} . By the closure condition on K , it follows that $\mathfrak{F} \in K$.

Suppose that \mathfrak{F} is not point-generated. But then as $\mathfrak{F} \Vdash \Lambda_K$, so does each point-generated subframe of \mathfrak{F} , hence all these subframes belongs to K . \mathfrak{F} is a bounded morphic image of the disjoint union of its rooted generated subframes, \square

3.4.2 The finite frame property

Definition 3.15. A normal modal logic Λ has the finite model property w.r.t. some class of models M if $M \Vdash \Lambda$ and every formula **not** in Λ is refuted in a **finite** model \mathfrak{M} in M . Λ has the finite model property if it has the finite model property w.r.t. some class of models

Informally, if a normal modal logic has the finite model property, it has a finite semantic characterization: **it is precisely the set of formulas that some collection of finite models makes globally true**

Definition 3.16 (Finite Frame Property). Let Λ be a normal modal logic and F a class of finite frames. We say Λ has the **finite frame property w.r.t.** F if and only if $F \Vdash \Lambda$, and for every formula ϕ s.t. $\phi \notin \Lambda$ there is some $\mathfrak{F} \in F$ s.t. ϕ is falsifiable on \mathfrak{F} . We say Λ has the **finite frame property** iff it has the finite frame property w.r.t. some class of finite frames

Definition 3.17 (Definable Variant). Let $\mathfrak{M} = (W, R, V)$ be a model and $U \subseteq W$. We say U is **definable in** \mathfrak{M} iff there is a formula ϕ_U s.t. for all states $w \in W$, $\mathfrak{M}, w \models \phi_U$ iff $w \in U$

Any model \mathfrak{M}' based on the frame (W, R) is called a **variant** of \mathfrak{M} . A variant (W, R, V') of \mathfrak{M} is **definable in** \mathfrak{M} iff for all proposition symbols p , $V'(p)$ is definable in \mathfrak{M} . If \mathfrak{M}' is a variant of \mathfrak{M} that is definable in \mathfrak{M} , we call \mathfrak{M}' a **definable variant** of \mathfrak{M}

Lemma 3.18. Let $\mathfrak{M} = (\mathfrak{F}, V)$ be a model and $\mathfrak{M}' = (\mathfrak{F}, V')$ be a definable variant of \mathfrak{M} . For any formula ϕ , let ϕ' be the result of uniformly replacing each atomic symbol p in ϕ by $\phi_{V'(p)}$, where $\phi_{V'(p)}$ defines $V'(p)$ in \mathfrak{M} . Then for all formulas ϕ , and all normal modal logics Λ :

1. $\mathfrak{M}', w \models \phi$ iff $\mathfrak{M}, w \models \phi'$
2. If every substitution instance of ϕ is true in \mathfrak{M} , then every substitution instance of ϕ is true in \mathfrak{M}'
3. If $\mathfrak{M} \models \Lambda$ then $\mathfrak{M}' \models \Lambda$

Proof. 1. Induction on ϕ . For the base case we have $\mathfrak{M}', w \models p$ iff $\mathfrak{M}, w \models \phi_{V'(p)}$. And $(\neg\phi)' = \neg\phi'$, $(\phi \vee \psi)' = \phi' \vee \psi'$, and $(\Diamond\phi)' = \Diamond\phi'$

2. Let ψ be a substitution instance of ϕ and suppose that $\mathfrak{M}' \not\models \psi$. Thus there is some w in \mathfrak{M} s.t. $\mathfrak{M}', w \not\models \psi$. By 1, this means $\mathfrak{M}, w \not\models \psi'$, which means that $\mathfrak{M} \not\models \psi'$. But as ψ' is a substitution instance of ψ and ψ is a substitution instance of ϕ , we have that ψ' is a substitution instance of ϕ

3. Normal modal logic are closed under uniform substitution

□

Definition 3.19 (Distinguishing Model). A model \mathfrak{M} is **distinguishing** if the relation \leftrightarrow of modal equivalence between states of \mathfrak{M} is the identity relation

Lemma 3.20. Let $\mathfrak{M} = (\mathfrak{F}, V)$ be a finite distinguishing model. Then

1. For every state w in \mathfrak{M} there is a formula ϕ_w that is true at and only at w
2. \mathfrak{M} can define any subsets of \mathfrak{F} . Hence \mathfrak{M} can define all its variants
3. If $\mathfrak{M} \models \phi$ then $\mathfrak{F} \models \phi$

Proof. 3. Suppose $\mathfrak{M} \Vdash \phi$, then we have $\mathfrak{M}' \Vdash \phi$ for some definable variant of \mathfrak{M} . □

Theorem 3.21. *A normal modal logic has the finite frame property iff it has the finite model property*

Proof. Suppose that Λ is a normal modal logic with the finite model property. and assume that $\Lambda \subset \text{Form}(\tau, \Phi)$

Take a formula in the language $ML(\tau, \Phi)$ that does not belong to Λ . We will show that ϕ can be refuted on a finite frame \mathfrak{F} s.t. $\mathfrak{F} \Vdash \Lambda$

As Λ has the finite model property, there is a finite model \mathfrak{M} s.t. $\mathfrak{M} \Vdash \Lambda$ and $\mathfrak{M}, w \not\Vdash \phi$ for some state $w \in \mathfrak{M}$. Let Σ be the set of all subformulas of formulas in $\{\phi\} \cup \Lambda$ and let \mathfrak{M}^f be any filtration of \mathfrak{M} through Σ . As \mathfrak{M} is finite, so is \mathfrak{M}^f . As \mathfrak{M}^f is a filtration, it is a distinguishing model. By the Filtration Theorem 2.25, $\mathfrak{M}^f, |w| \not\Vdash \phi$. Moreover $\mathfrak{M}^f \Vdash \Lambda$. Let \mathfrak{F} be the (finite) frame underlying \mathfrak{M}^f . By Lemma ?? $\mathfrak{F} \Vdash \Lambda$ □

Exercise 3.4.1. Let \mathfrak{F} be a transitive, finite, point-generated frame, let w be a root of \mathfrak{F} , and let $\phi_{\mathfrak{F}, w}$ be the Jankov-Fine formula for \mathfrak{F} and w . Then for any transitive frame \mathfrak{G} we have the following equivalence: there is a valuation V and a node v s.t. $(\mathfrak{G}, V), v \Vdash \phi_{\mathfrak{F}, w}$ iff there exists a bounded morphism from \mathfrak{G}_v onto \mathfrak{F}

Proof. Left to right. Suppose there is a valuation V and a node v s.t. $(\mathfrak{G}, V), v \Vdash \phi_{\mathfrak{F}, w}$. 1 shows that $v \in V(p_0)$. 2 and 3 shows that each of states satisfies exactly one of proposition letter p_i . Hence we build a function $f : \mathfrak{G}_v \rightarrow \mathfrak{F}$ as $f(v_i) = w_j$ if $v_i \in V(p_j)$ for each $v_i \in \mathfrak{G}$. 2 and 3 show that this function is well-defined.

Suppose $R_{\mathfrak{G}} v_i v_j$ and $\neg R_{\mathfrak{F}} w_i w_j$, then $(\mathfrak{G}, V), g \Vdash p_i \rightarrow \neg \Diamond p_j$ for all $g \in \mathfrak{G}$. However, $(\mathfrak{G}, V), v_i \Vdash p_i$ and $(\mathfrak{G}, V), v_j \Vdash p_j$

If $R_{\mathfrak{F}} f(v_i) w_j$, then $(\mathfrak{G}, V), v_i \Vdash p_{f(v_i)} \rightarrow \Diamond p_j$. Hence there is a successor v_j of v_i s.t. $f(v_j) = w_j$. Therefore we find $R_{\mathfrak{G}} v_i v_j$ and f is a bounded morphism

Right to left. Suppose there is a bounded morphism f from \mathfrak{G}_v onto \mathfrak{F} . Then clearly $f(v) = w_0$. Let $V(p_{f(v_i)}) = \{v_j \mid f(v_j) = f(v_i)\}$ for all $v_i \in \mathfrak{G}$. Then $(\mathfrak{G}, V), v \Vdash p_0 \wedge \Box(p_0 \vee \dots \vee p_n) \wedge (p_i \rightarrow \neg p_j) \wedge \Box(p_i \rightarrow \neg p_j)$. If $R_{\mathfrak{F}} w_i w_j$ and there is a $v_i \in \mathfrak{G}$ s.t. $f(v_i) = w_i$, then there is a $f(v_j) = w_j$ s.t. $R_{\mathfrak{G}} v_i v_j$. Hence $(\mathfrak{F}, V), v \Vdash (p_i \rightarrow \Diamond p_j) \wedge \Box(p_i \rightarrow \Diamond p_j)$ for each i, j with $R w_i w_j$. 5 is the same □

Exercise 3.4.2. Let \mathfrak{M} be a model, let \mathfrak{M}^f be any filtration of \mathfrak{M} through some finite set of formulas Σ , and let f be the natural map associated with the filtration. If u is a point in the filtration, show that $f^{-1}[u]$ is definable in \mathfrak{M}

Proof. $f^{-1}[u] = \{w \mid w \leftrightarrow_{\Sigma} u\}$. Since \mathfrak{M}^f is distinguishing, there is a formula ϕ_u that is true at, and only at, u . Then $\mathfrak{M}, w \models \phi_u$ iff $w \in f^{-1}[u]$ \square

4 Completeness

4.1 Preliminaries

Definition 4.1 (Modal Logics). A **modal logic** Λ is a set of modal formulas that contains all propositional tautologies and is closed under **modus ponens** and **uniform substitution**. If $\phi \in \Lambda$ we say that ϕ is a **theorem** of Λ and write $\vdash_{\Lambda} \phi$. If Λ_1 and Λ_2 are modal logics s.t. $\Lambda_1 \subseteq \Lambda_2$, we say that Λ_2 is an **extension** of Λ_1

Example 4.1. 3. Define Λ_S be to $\{\phi \mid \mathfrak{G} \models \phi, \text{ for all structures } \mathfrak{G} \in S\}$, where S is any class of frames or any class of general frames. Λ_S is a logic.

4. If M is a class of models, then Λ_M need **not** be a logic. Consider a model \mathfrak{M} in which p is true at all nodes but q is not. Then $p \in \Lambda_{\mathfrak{M}}$ but $q \notin \Lambda_{\mathfrak{M}}$. But q is obtainable from p by uniform substitution

The logic generated by the empty set contains all the tautologies and nothing else; we call it **PC**

Definition 4.2. Let $\psi_1, \dots, \psi_n, \phi$ be modal formulas. We say that ϕ is **deducible in propositional calculus from assumptions** ψ_1, \dots, ψ_n if $(\psi_1 \wedge \dots \wedge \psi_n) \rightarrow \phi$ is a tautology

Definition 4.3. If $\Gamma \cup \{\phi\}$ is a set of formulas then ϕ is **deducible in Λ from Γ** if $\vdash_{\Lambda} \phi$ or there are formulas $\psi_1, \dots, \psi_n \in \Gamma$ s.t.

$$\vdash_{\Lambda} (\psi_1 \wedge \dots \wedge \psi_n) \rightarrow \phi$$

A set of formulas Γ is **Λ -consistent** if $\Gamma \not\vdash_{\Lambda} \perp$.

See more in Does the deduction theorem fail for modal logic? and MO discussion

Definition 4.4. A modal logic Λ is **normal** if it contains the formulas

$$\begin{aligned} \text{(K)} \quad & \Box(p \rightarrow q) \rightarrow (\Box p \rightarrow \Box q) \\ \text{(Dual)} \quad & \Diamond \leftrightarrow \neg \Box \neg p \end{aligned}$$

and is closed under generalization (if $\vdash_{\Lambda} \phi$ then $\vdash_{\Lambda} \Box \phi$)

Lemma 4.5. For any normal logic Λ , if $\vdash_{\Lambda} \phi \leftrightarrow \psi$ then $\vdash_{\Lambda} \Diamond \phi \leftrightarrow \Diamond \psi$

The normal logic generated by the empty set is called **K**. If Γ is a non-empty set of formulas we usually denote the normal logic generated by Γ by **K** Γ

$$\begin{aligned} (4) \quad & \Diamond \Diamond p \rightarrow \Diamond p \\ (T) \quad & p \rightarrow \Diamond p \\ (B) \quad & p \rightarrow \Box \Diamond p \\ (D) \quad & \Box p \rightarrow \Diamond p \\ (.3) \quad & \Diamond p \wedge \Diamond q \rightarrow \Diamond(p \wedge \Diamond q) \vee \Diamond(p \wedge q) \vee \Diamond(q \wedge \Diamond p) \\ (L) \quad & \Box(\Box p \rightarrow p) \rightarrow \Box p \end{aligned}$$

K	the class of all frames
K4	the class of transitive frames
T	the class of reflexive frames
B	the class of symmetric frames
KD	the class of right-unbounded frames
S4	the class of reflexive, transitive frames
S5	the class of frames whose relation is an equivalence relation
K4.3	the class of transitive frames with no branching to the right
S4.3	the class of reflexive, transitive frames with no branching to the right
KL	the class of finite transitive trees

A frame (W, R) satisfying $\forall x \forall y \forall z (Rxy \wedge Rxz \rightarrow (Ryz \vee y = z \vee Rzy))$ is said to have **no branching to the right**

Definition 4.6 (Soundness). Let S be a class of frames. A normal modal logic Λ is **sound** w.r.t. S if $\Lambda \subseteq \Lambda_S$. (Equivalently: Λ is **sound** w.r.t. S if for all formulas ϕ , and all structures $\mathfrak{G} \in S$, $\vdash_{\Lambda} \phi$ implies $\mathfrak{G} \models \phi$). If Λ is sound w.r.t. S we say that S is a **class of frames for Λ**

Definition 4.7 (Completeness). Let S be a class of frames. A logic Λ is **strongly complete** w.r.t. S if for any set of formulas $\Gamma \cup \{\phi\}$, if $\Gamma \Vdash_S \phi$ then $\Gamma \vdash_\Lambda \phi$.

A logic Λ is **weakly complete** w.r.t. S if for any formula ϕ , if $S \models \phi$ then $\vdash_\Lambda \phi$.

A weakly complete w.r.t. S if $\Lambda_S \subseteq \Lambda$. Thus if we prove that a syntactically specified logic Λ is both sound and weakly complete w.r.t. some class of structures S , we have $\Lambda = \Lambda_S$.

Proposition 4.8. *A logic Λ is strongly complete w.r.t. a class of structures S iff every Λ -consistent set of formulas is satisfiable on some $\mathfrak{S} \in S$. Λ is weakly complete w.r.t. a class of structures S iff every Λ -consistent formula is satisfiable on some $\mathfrak{S} \in S$.*

Proof. Suppose Λ is not strongly complete w.r.t. S . There is a set of formulas $\Gamma \cup \{\phi\}$ s.t. $\Gamma \Vdash_S \phi$ but $\Gamma \not\vdash_\Lambda \phi$. Then $\Gamma \cup \{\neg\phi\}$ is Λ -consistent but not satisfiable on any structure in S .

Suppose there is a Λ -consistent set Γ of formulas is not satisfiable on all $\mathfrak{S} \in S$. Then $\Gamma \Vdash_S \phi$ is always true \square

Definition 4.9. Assume we are working with a modal language of similarity type τ . A **modal logic** in this language is a set of formulas containing all tautologies that is closed under modus ponens and uniform substitution. A modal logic Λ is **normal** if for every operator ∇ it contains the axiom K_∇^i (for all i s.t. $1 \leq i \leq \rho(\nabla)$) and the axiom Dual_∇ and is closed under the generalization rules

$$\begin{aligned} (K_\nabla^i) \quad & \nabla(r_1, \dots, p \rightarrow q, \dots, r_{\rho(\nabla)}) \\ & \rightarrow (\nabla(r_1, \dots, p, \dots, r_{\rho(\nabla)}) \rightarrow \nabla(r_1, \dots, q, \dots, r_{\rho(\nabla)})) \\ (\text{Dual}_\nabla) \quad & (r_1, \dots, r_{\rho(\nabla)}) \leftrightarrow \neg \nabla(\neg r_1, \dots, \neg r_{\rho(\nabla)}) \end{aligned}$$

for a polyadic operator ∇ , generalization takes the following form

$$\vdash_\Lambda \sigma \text{ implies } \vdash_\Lambda \nabla(\perp, \dots, \sigma, \dots, \perp)$$

Definition 4.10. Let τ be a modal similarity type. Given a set of τ -formulas Γ , we define $K_\tau \Gamma$, the normal modal logic **axiomatized** or **generated** by Γ , to be the smallest normal modal τ -logic containing all formulas in Γ . Formulas in Γ are called **axioms** of this logic, and Γ may be called an **axiomatization** of $K_\tau \Gamma$.

4.2 Canonical Models

Definition 4.11 (Λ -MCSs). A set of formulas Γ is **maximal Λ -consistent** if Γ is Λ -consistent, and any set of formulas properly containing Γ is Λ -inconsistent. We say its a Λ -MCS

Proposition 4.12 (Properties of MCSs). *If Λ is a logic and Γ is a Λ -MCS then*

1. Γ is closed under modus ponens: if $\phi, \phi \rightarrow \psi \in \Gamma$, then $\psi \in \Gamma$
2. $\Lambda \subseteq \Gamma$
3. for all formulas ϕ : $\phi \in \Gamma$ or $\neg\phi \in \Gamma$
4. for all formulas ϕ, ψ : $\phi \vee \psi \in \Gamma$ iff $\psi \in \Gamma$ or $\phi \in \Gamma$

Proposition 4.13 (Lindenbaum's Lemma). *If Σ is a Λ -consistent set of formulas then there is a Λ -MCS Σ^+ s.t. $\Sigma \subseteq \Sigma^+$*

Proof. Let ϕ_0, ϕ_1, \dots be an enumeration of the formulas of our language. We define the set Σ^+ as the union of a chain of Λ -consistent sets as follows:

$$\begin{aligned}\Sigma_0 &= \Sigma \\ \Sigma_{n+1} &= \begin{cases} \Sigma_n \cup \{\phi_n\} & \text{if this is } \Lambda\text{-consistent} \\ \Sigma_n \cup \{\neg\phi_n\} & \text{otherwise} \end{cases} \\ \Sigma^+ &= \bigcup_{n \geq 0} \Sigma_n\end{aligned}$$

1. Σ_n is Λ -consistent, for all n
2. exactly one of ϕ and $\neg\phi$ is in Σ^+
3. if $\Sigma^+ \vdash_{\Lambda} \phi$, then $\phi \in \Sigma^+$
4. Σ^+ is a Λ -MCS

□

Definition 4.14. The **canonical model** \mathfrak{M}^{Λ} for a normal modal logic Λ is the triple $(W^{\Lambda}, R^{\Lambda}, V^{\Lambda})$ where

1. W^{Λ} is the set of all Λ -MCSs
2. R^{Λ} is the binary relation on W^{Λ} defined by $R^{\Lambda}wu$ if for all formulas ψ , $\psi \in u$ implies $\Diamond\psi \in w$. R^{Λ} is called the **canonical relation**

3. V^Λ is the valuation defined by $V^\Lambda(p) = \{w \in W^\Lambda \mid p \in w\}$. V^Λ is called the **canonical valuation**

Lemma 4.15. *For any normal logic Λ , $R^\Lambda wv$ iff for all formulas ψ , $\Box\psi \in w$ implies $\psi \in v$*

Proof. Suppose $R^\Lambda wv$ and $\psi \notin v$. As v is an MCS, $\neg\psi \in v$. The $\Diamond\neg\psi \in w$. As w is consistent, $\neg\Diamond\neg\psi \notin w$. That is, $\Box\psi \notin w$ \square

Lemma 4.16 (Existence Lemma). *For any normal modal logic Λ and any state $w \in W^\Lambda$, if $\Diamond\phi \in w$ then there is a state $v \in W^\Lambda$ s.t. $R^\Lambda wv$ and $\phi \in v$*

Proof. Suppose $\Diamond\phi \in w$. Let $v^- = \{\phi\} \cup \{\psi \mid \Box\psi \in w\}$. Then v^- is consistent. For suppose not. Then there are ψ_1, \dots, ψ_n s.t. $\vdash_\Lambda (\psi_1 \wedge \dots \wedge \psi_n) \rightarrow \neg\phi$, then we have $\vdash_\Lambda \Box(\psi_1 \wedge \dots \wedge \psi_n) \rightarrow \Box\neg\phi$. Also as $(\Box\psi_1 \wedge \dots \wedge \Box\psi_n) \rightarrow \Box(\psi_1 \wedge \dots \wedge \psi_n)$, then we $\vdash_\Lambda (\Box\psi_1 \wedge \dots \wedge \Box\psi_n) \rightarrow \Box\neg\phi$. Now $\Box\psi_1 \wedge \dots \wedge \Box\psi_n \in w$ thus it follows that $\Box\neg\phi \in w$. Using dual, it follows that $\neg\Diamond\phi \in w$, which is impossible.

Let v be any MCS extending v^- ; such extensions exist by Lindenbaum's Lemma. By construction $\phi \in v$. Furthermore, for all formulas ψ , $\Box\psi \in w$ implies $\psi \in v$. Hence by Lemma 4.15, $R^\Lambda wv$ \square

Lemma 4.17 (Truth Lemma). *For any normal modal logic Λ and any formula ϕ , $\mathfrak{M}^\Lambda, w \Vdash \phi$ iff $\phi \in w$*

Proof. Base case follows from the definition of V^Λ . Boolean case from Proposition 4.12.

$$\begin{aligned} \mathfrak{M}^\Lambda, w \Vdash \Diamond\phi &\Leftrightarrow \exists v(R^\Lambda wv \wedge \mathfrak{M}^\Lambda, v \Vdash \phi) \\ &\Leftrightarrow \exists v(R^\Lambda wv \wedge \phi \in v) \\ &\Rightarrow \Diamond\phi \in w \end{aligned}$$

Suppose $\Diamond\phi \in w$, it suffices to find an MCS v s.t. $R^\Lambda wv$ and $\phi \in v$, which is the existence lemma \square

Theorem 4.18 (Canonical Model Theorem). *Any normal modal logic is strongly complete w.r.t. its canonical model*

Proof. Suppose Σ is a consistent set of the normal modal logic Λ . By Lindenbaum's Lemma there is a Λ -MCS Σ^+ extending Σ . By the previous lemma, $\mathfrak{M}^\Lambda, \Sigma^+ \Vdash \Sigma$ \square

Theorem 4.19. *K is strongly complete w.r.t. to the class of all frames*

. By Proposition 4.8, to prove this result it suffices to find, for any **K**-consistent set of formulas Γ , a model \mathfrak{M} and a state w in \mathfrak{M} s.t. $\mathfrak{M}, w \Vdash \Gamma$. Simply choose \mathfrak{M} to be $(\mathfrak{F}^{\mathbf{K}}, V^{\mathbf{K}})$ \square

4.3 Applications

Theorem 4.20. *The logic K4 is strongly complete w.r.t. the class of transitive frames*

Proof. Given a **K4**-consistent set of formulas Γ , it suffices to find a model (\mathfrak{F}, V) and a state w in this model s.t.

1. $(\mathfrak{F}, V), w \Vdash \Gamma$ and
2. \mathfrak{F} is transitive

$(W^{\mathbf{K4}}, R^{\mathbf{K4}}, V^{\mathbf{K4}}), \Gamma^+ \Vdash \Gamma$. Suppose $R^{\mathbf{K4}} wv$ and $R^{\mathbf{K4}} vu$. We wish to show that $R^{\mathbf{K4}} wu$. Suppose $\phi \in u$, then $\Diamond\phi \in v$, so $\Diamond\Diamond\phi \in w$. But w is a **K4**-MCS, hence it contains $\Diamond\Diamond\phi \rightarrow \Diamond\phi$ \square

Theorem 4.21. *T, KB, KD are strongly complete w.r.t. the classes of reflexive frames, of symmetric frames, and of right-unbounded frames respectively*

Theorem 4.22. *S4 is strongly complete w.r.t. the class of reflexive, transitive frames. S5 is strongly complete w.r.t. the class of frames whose relation is an equivalence relation*

Definition 4.23 (Canonicity). A formula ϕ is **canonical** if for any normal logic Λ , $\phi \in \Lambda$ implies that ϕ is valid on the canonical frame for Λ . A normal logic Λ is **canonical** if its canonical frame is a frame for Λ . (That is, Λ is canonical if for all ϕ s.t. $\vdash_{\Lambda} \phi$, ϕ is valid on the canonical frame for Λ)

Λ is strongly complete w.r.t. canonical
4, T, B, D axioms are all canonical formulas.

Definition 4.24 (Canonicity for a Property). Let ϕ be a formula, and P be a property. If the canonical frame for any normal logic Λ containing ϕ has property P , then ϕ is valid on any class of frames with property P , then ϕ is **canonical for P**.

The class of all bidirectional frames is denoted by F_t , and a bidirectional model is a model whose underlying frame belongs to F_t

Definition 4.25. The minimal temporal logic Λ_t is $\{\phi \mid F_t \Vdash \phi\}$

Definition 4.26. A **normal temporal logic** Λ is a normal modal logic that contains $p \rightarrow GPs$ and $p \rightarrow HFs$. The smallest normal temporal logic is called \mathbf{K}_t . We usually call normal temporal logics **tense logics**

We want to show that \mathbf{K}_t generated exactly the formulas in Λ_{F_t} .

Definition 4.27. The canonical model for a tense logic Λ is the structure $\mathfrak{M}^\Lambda = (T^\Lambda, \{R_P^\Lambda, R_F^\Lambda\}, V^\Lambda)$ where

1. T^Λ is the set of all Λ -MCSs
2. R_P^Λ is the binary relation on T^Λ defined by $R_P^\Lambda ts$ if for all formulas ϕ , $\phi \in s$ implies $P\phi \in t$
3. R_F^Λ is the binary relation on T^Λ defined by $R_F^\Lambda ts$ if for all formulas ϕ , $\phi \in s$ implies $F\phi \in t$
4. V^Λ is the valuation defined by $V^\Lambda(p) = \{t \in T^\Lambda \mid p \in t\}$

Lemma 4.28. For any tense logic Λ , if $R_P^\Lambda ts$ then $R_F^\Lambda st$ and if $R_F^\Lambda ts$ then $R_P^\Lambda st$

Proof. If $R_P^\Lambda ts$ and suppose $\phi \in t$. Then $GP\phi \in t$. Hence $P\phi \in s$ and we have $R_F^\Lambda st$ \square

Corollary 4.29. \mathbf{K}_t is strongly complete w.r.t. the class of all bidirectional frames, and $\mathbf{K}_t = \Lambda_{F_t}$

Proof. \mathbf{K}_t is strongly complete w.r.t. its canonical model. As we have just seen, this model is based on a *bidirectional* frame, so the strong frame completeness result follows. Strong completeness implies weak completeness, so $\Lambda_{F_t} \subseteq \mathbf{K}_t$ \square

Definition 4.30. A bidirectional frame (T, R) is **dense** if there is a point between any two related points ($\forall xy(Rxy \rightarrow \exists z(Rxz \wedge Rzy))$). It is **right-unbounded** if every points has a successor. **left-unbounded**. **unbounded** if both right- and left-unbounded. It is **trichotomous** if any two points are equal or are related one way or the other ($\forall xy(Rxy \vee x = y \vee Ryx)$) and a **weak total order** if it is both transitive and trichotomous. We call a frame with all these properties a DUWTO-frame

$$\begin{aligned}
(4) \quad & FFp \rightarrow Fp \\
(D_r) \quad & Gp \rightarrow Fp \\
(D_l) \quad & Hp \rightarrow Pp
\end{aligned}$$

We obtain an unbounded canonical frame.

$$(den) \quad Fp \rightarrow FFp$$

Lemma 4.31. $Fp \rightarrow FFp$ is canonical for density

Proof. Let Λ be any tense logic containing $Fp \rightarrow FFp$, let (T^Λ, R^Λ) be its canonical frame, and let t, t' be points in this frame s.t. $R^\Lambda tt'$. We have to show that there is a Λ -MCS s s.t. $R^\Lambda ts$ and $R^\Lambda st'$. If we could show that $\{\phi \mid G\phi \in t\} \cup \{F\psi \mid \psi \in t'\}$ is consistent we would have the desired result by Lemma 4.15 and 4.28

So suppose for the sake of contradiction that this set is not consistent. Then, for some finite set of formulas $\phi_1, \dots, \phi_m, \psi_1, \dots, \psi_n$ from this set

$$\vdash_\Lambda (\phi_1 \wedge \dots \wedge \phi_m \wedge F\psi_1 \wedge \dots \wedge F\psi_n) \rightarrow \perp$$

Define $\hat{\phi}$ to be $\phi_1 \wedge \dots \wedge \phi_m$ and $\hat{\psi}$ to be $\psi_1 \wedge \dots \wedge \psi_n$. Note that $\hat{\psi} \in t'$.

Now $\vdash_\Lambda F\hat{\psi} \rightarrow F\psi_1 \wedge \dots \wedge F\psi_n$, hence $\vdash_\Lambda \hat{\phi} \wedge F\hat{\psi} \rightarrow \perp$, hence $\vdash_\Lambda \hat{\phi} \rightarrow \neg F\hat{\psi}$, and hence $\vdash_\Lambda G\hat{\phi} \rightarrow G\neg F\hat{\psi}$. Because $G\phi_1, \dots, G\phi_m \in t$, we have that $G\hat{\phi} \in t$ too, hence $G\neg F\hat{\psi} \in t$, and hence $\neg G\neg F\hat{\psi} \notin t$. That is $FF\hat{\psi} \notin t$. But this means that $F\hat{\psi} \notin t$ as $F\hat{\psi} \rightarrow FF\hat{\psi} \in t$. But now we have a contradiction: as $\hat{\psi} \in t'$ and $R^\Lambda tt', F\hat{\psi} \in t$ □

$$\begin{aligned}
(.3_r) \quad & (Fp \wedge Fq) \rightarrow F(p \wedge Fq) \vee F(p \wedge q) \vee F(q \wedge Fp) \\
(.3_l) \quad & (Pp \wedge Pq) \rightarrow P(p \wedge Pq) \vee P(p \wedge q) \vee P(q \wedge Pp)
\end{aligned}$$

which is canonical for no-branching-to-the -left and -right respectively

Proposition 4.32. Any trichotomous frame (T, R) is non-branching. Furthermore, if R is transitive and non-branching and $t \in T$, then the subframe of (T, R) generated by t is trichotomous

Definition 4.33. Let $\mathbf{K}_t\mathbf{Q}$ be the smallest tense logic containing 4, D_l , D_r , den , $.3_l$ and $.3_r$

Theorem 4.34. K_tQ is strongly complete w.r.t. the class of DUWTO-frames

Proof. If Γ is a K_tQ -consistent set of formulas, extend it to a K_tQ -MCS Γ^+ . Let \mathfrak{M} be the canonical model for K_tQ and let fG be the submodel of \mathfrak{M} generated by Γ^+ . Then $\mathfrak{G}, \Gamma^+ \Vdash \Gamma$. Moreover, the frame underlying \mathfrak{G} is a DUWTO-frame as required. Conjunction of $.3_l$ and $.3_r$ is canonical for non-branching, \mathfrak{M} is non-branching and \mathfrak{G} trichotomous \square

Theorem 4.35. Every Sahlqvist formula is canonical for the first-order property it defines. Hence given a set of Sahlqvist axioms Σ , the logic K is strongly complete w.r.t. the class of frames

Exercise 4.3.1. Let 1.1 be the axiom $\Diamond p \rightarrow \Box p$. Show that **K1.1** is sound and strongly complete w.r.t. the class of all frames (W, R) s.t. R is a partial function

Proof. Suppose $\mathfrak{F} = (W, R)$ is a frame s.t. R is a partial function and $\mathfrak{F} \Vdash \Diamond p$. Then for an arbitrary valuation V and a point w , we have $(\mathfrak{F}, V), w \Vdash \Diamond p$. Then there is a point v s.t. Rwv and $v \in V(p)$. As R is a partial function, w has no successor other than v . Hence $(\mathfrak{F}, V), w \Vdash \Box p$. Hence $\mathfrak{F} \Vdash \Diamond p \rightarrow \Box p$.

Let $\mathfrak{M}^C = (M^C, R^C, V^C)$ be the canonical model of **K1.1**. It suffices to prove R^C is a partial function. Suppose there are states w, u, v s.t. R^Cwu and R^Cwv . For any $\phi \in u$, $\Diamond \phi \in w$ and $\mathfrak{M}^C, w \Vdash \Diamond \phi \rightarrow \Box \phi$. Hence $\mathfrak{M}^C, w \Vdash \Box \phi$ and $\phi \in v$ follows. Hence $u \subseteq v$ and $v \subseteq u$ the same. So $u = v$ and R is a partial function. \square

Exercise 4.3.2. Let Λ be a normal temporal logic containing the axioms $p \rightarrow GPP$ and $p \rightarrow HFP$. Show that if $R_P^\Lambda ts$ then $R_F^\Lambda st$ and if $R_F^\Lambda ts$ then $R_P^\Lambda st$

Proof. If $R_P^\Lambda ts$ and suppose $\phi \in t$. Then $GP\phi \in t$. Hence $P\phi \in s$ and we have $R_F^\Lambda st$ \square

Exercise 4.3.3. Show that **S5** is complete w.r.t. the class of **globally related frames**, that is, those frames (W, R) where R connects any two points

Proof. Relation in canonical model of **S5** is equivalent. Given any **S5**-consistent set of formulas Γ , $\mathfrak{M}^{S5}, \Gamma^+ \Vdash \Gamma$. Consider the submodel \mathfrak{G} of canonical model generated on Γ^+ , clearly $\mathfrak{G}, \Gamma^+ \Vdash \Gamma$. Obviously the elements in \mathfrak{G} is the equivalence class of Γ^+ , hence the frame underlying \mathfrak{G} is globally related. \square

Exercise 4.3.4. Use canonical models to show that **K4.3** is strongly complete w.r.t. the class of frames that are transitive and have no branching to the right. Then, by proving suitable completeness results, show that the normal logic axiomatized by 4 and $\Box(p \wedge \Box p \rightarrow q) \vee \Box(q \wedge \Box q \rightarrow p)$ is **K4.3**

Proof. Let $\mathfrak{M}^C = (W^C, R^C, V^C)$ be a canonical model of **K4.3**. It suffices to show that R^C is transitive and have no branching to the right. Suppose R^C_{wv} and R^C_{wu} . \square

4.4 Limitative Results

KL is the normal modal logic generated by the Löb axiom $\Box(\Box p \rightarrow p) \rightarrow \Box p$.

Every **canonical** logic is sound and strongly complete w.r.t. some class of frames

Theorem 4.36. *KL is not sound and strongly complete w.r.t. any class of frames, and hence it is not canonical*

Proof. Let Γ be $\{\Diamond q_1\} \cup \{\Box(q_i \rightarrow \Diamond q_{i+1}) \mid 1 \leq i \in \omega\}$. We will show that Γ is **KL**-consistent, and that no model based on a **KL**-frame can satisfy all formulas in Γ at a single point.

To show that Γ is consistent, it suffices to show that every finite subset Ψ of Γ is consistent. Given any such Ψ , for some natural number n there is a finite set Φ of the form $\{\Diamond q_1\} \cup \{\Box(q_i \rightarrow \Diamond q_{i+1}) \mid 1 \leq i < n\}$ s.t. $\Psi \subseteq \Phi \subset \Gamma$. We show that Φ , and hence Ψ , is consistent.

Let $\widehat{\Phi}$ be the conjunction of all the formulas in Φ . To show that $\widehat{\Phi}$ is **KL**-consistent, it suffices to show that it can be satisfied in a model based on a frame for **KL**, for this shows that $\neg \widehat{\Phi}$ is not valid on all frames for **KL**. Let \mathfrak{F} be the frame consisting of $\{0, \dots, n\}$ in their usual order; as this is a transitive, converse well-founded frame by Example 3.4 it is a frame for **KL**. Let \mathfrak{M} be any model based on \mathfrak{F} s.t. for all $1 \leq i \leq n$, $V(q_i) = \{i\}$. Then $\mathfrak{M}, 0 \Vdash \widehat{\Phi}$ and $\widehat{\Phi}$ is **KL**-consistent

Next, suppose that **KL** is sound and strongly complete w.r.t. some class of frames F ; note that as **KL** is not the inconsistent logic, F must be non-empty. Thus any **KL**-consistent set of formulas can be satisfied at some point in a model based on a frame in F . In particular, there is $\mathfrak{M}, w \Vdash \Gamma$. But this is impossible: because $\mathfrak{M}, w \Vdash \Gamma$ we inductively define an infinite path through \mathfrak{M} starting at w . \square

Remark. A normal logic Λ is said to be **compact** when any Λ -consistent set Σ can be satisfied in a frame for Λ at a single point. A non-compact logic cannot be canonical, sound and strongly compact

Theorem 4.37. *KL is weakly complete w.r.t. the class of all finite transitive trees*

Definition 4.38. Let Λ be a normal modal logic. Λ is **(frame) complete** if there is a class of frames \mathcal{F} s.t. $\Lambda = \Lambda_{\mathcal{F}}$ and **(frame) incomplete**

$\mathbf{K}_t\mathbf{Tho}$ is the tense logic generated by the following axioms

$$\begin{aligned} (.3_r) & Fp \wedge Fq \rightarrow F(p \wedge Fq) \vee F(p \wedge q) \vee F(Fp \wedge q) \\ (D_r) & Gp \rightarrow Fp \\ (L_l) & H(Hp \rightarrow p) \rightarrow Hp \end{aligned}$$

The first two axioms are no branching to the right and right-unboundedness respectively. The third is the Löb axiom written in terms of the backwards. As all three axioms are valid on the natural numbers, $\mathbf{K}_t\mathbf{Tho}$ is consistent. If (T, R) is a frame for $\mathbf{K}_t\mathbf{Tho}$ and $t \in T$, then $\{u \in T \mid Rtu\}$ is a right-unbounded strict total order

Let $\mathbf{K}_t\mathbf{ThoM}$ be the smallest tense logic containing $\mathbf{K}_t\mathbf{Tho}$ and the McK-insey axiom $GFp \rightarrow FGp$

Definition 4.39. Let $(U, <)$ be a strict total order and $S \subseteq U$. S is **cofinal in** U if for every $u \in U$ there is an $s \in S$ s.t. $u < s$

Lemma 4.40. *Let \mathfrak{F} be any frame for $\mathbf{K}_t\mathbf{Tho}$. Then $\mathfrak{F} \not\models GFp \rightarrow FGp$*

Proof. Let t be any point in \mathfrak{F} , let $U = \{u \in T \mid Rtu\}$ and let $<$ be the restriction of R to U . As \mathfrak{F} validates all the $\mathbf{K}_t\mathbf{Tho}$ axioms, $(U, <)$ is a right-unbounded strict order. Suppose we could show that there is a non-empty proper subset S of U s.t. both S and $U \setminus S$ are cofinal in U , then the lemma would be proved, for we would merely need to define a valuation V on \mathfrak{F} s.t. $V(p) = S$ and $(\mathfrak{F}, V), t \not\models GFp \rightarrow FGp$

Take a cardinal κ that is larger than the size of U . By ordinal induction, we will define a sequence of pairs of sets $(R_\alpha, S_\alpha)_{\alpha \leq \kappa}$ s.t. $R_\kappa \cap S_\kappa = \emptyset$ and both R_κ and S_κ are cofinal. We can prove the lemma from this by taking $S = S_\kappa$. The definition is as follows

1. For $\alpha = 0$, take some points r_0 and s_0 in U s.t. $r_0 < s_0$ and define $R_0 = \{r_0\}$ and $S_0 = \{s_0\}$
2. If $\alpha = \beta + 1$
 - (a) if R_β or S_β is cofinal, then define $R_\alpha = R_\beta$ and $S_\alpha = S_\beta$
 - (b) if neither R_β nor S_β is cofinal, then take some upper bound r_β of S_β take some s_β bigger than r_β and define $R_\alpha = R_\beta \cup \{r_\beta\}$ and $S_\alpha = S_\beta \cup \{s_\beta\}$



Figure 1: A model and its unraveling

(c) If α is a limit ordinal then define $R_\alpha = \bigcup_{\beta < \alpha} R_\beta$ and $S_\alpha = \bigcup_{\beta < \alpha} S_\beta$

It's easy to prove that $R_\alpha \cap S_\alpha = \emptyset$ for every ordinal $\alpha \leq \kappa$. If R_κ and S_κ were not cofinal, then the map $f : \kappa \rightarrow U$ given by $f : \alpha \mapsto r_{\alpha+1}$ would be injective \square

Theorem 4.41. *$K_t\text{ThoM}$ is consistent and incomplete*

Proof. Let $(\mathbb{N}, <)$ be the natural numbers in their usual order. Let A be the collection of finite and co-finite subsets of \mathbb{N} . A is closed under boolean combinations and modal operations. \square

4.5 Transforming the Canonical Model

Definition 4.42 (Unraveling). Let (W, R) be a frame generated by some point $w \in W$. The **unraveling** of (W, R) around w is the frame (\vec{W}, \vec{R}) where

1. \vec{W} is the set of finite sequences (w, w_1, \dots, w_n) s.t. $w_1, \dots, w_n \in W$ and $Rww_1, \dots, Rww_{n-1}w_n$
2. If $\vec{s}_1, \vec{s}_2 \in \vec{W}$, then $\vec{R}\vec{s}_1\vec{s}_2$ if there is some $v \in W$ s.t. $\vec{s}_1 + (v) = \vec{s}_2$ where $+$ denotes sequence concatenation

$$\vec{V}(p) = \{(w, w_1, \dots, w_n) \in \vec{W} \mid w_n \in V(p)\}$$

The model $\vec{\mathfrak{M}} = (\vec{W}, \vec{R}, \vec{V})$ is called the unraveling of \mathfrak{M} around w

Unraveling yields an irreflexive, intransitive and asymmetric frame.

Lemma 4.43. *Let $\vec{\mathfrak{M}} = (\vec{W}, \vec{R}, \vec{V})$ be the unraveling of $\mathfrak{M} = (W, R, V)$ around w . Then (W, R) is a bounded morphic image of (\vec{W}, \vec{R}) , and \mathfrak{M} is a bounded morphic image of $\vec{\mathfrak{M}}$*

Proof. Let $f : \vec{W} \rightarrow W$ be defined by $f(w, w_1, \dots, w_n) = w_n$. f is surjective, has the back and forth properties.

If $Rf(\vec{w})v$, then let $\vec{v} = \vec{w} + (v)$. Hence $Rf(\vec{w})f(\vec{v})$ \square

Any satisfiable set of formulas is satisfiable on a (irreflexive, intransitive and asymmetric) tree: for if a set of formulas is satisfiable, it is satisfiable on a point-generated model, hence by unraveling we have the result. It follows that **K** is (strongly) complete w.r.t. this class of models

Suppose we unravel \mathfrak{M} around some generating point w to obtain $(\vec{W}, \vec{R}, \vec{V})$. Now consider the model $\mathfrak{M}^* = (\vec{W}, R^*, \vec{V})$ where R^* is the reflexive closure of \vec{R} . Trivially \mathfrak{M}^* is an **S4** model. Moreover, as (\vec{W}, \vec{R}) is a tree, (\vec{W}, R^*) is an *antisymmetry* frame.

Lemma 4.44. Let $\mathfrak{M} = (W, R, V)$ be a reflexive transitive model generated by some $w \in W$, and let $(\vec{W}, \vec{R}, \vec{V})$ be the unraveling of \mathfrak{M} around w . Let R^* be the reflexive transitive closure of \vec{R} , and define \mathfrak{M}^* to be (\vec{W}, R^*, \vec{V}) . Then \mathfrak{M} is a bounded morphic image of \mathfrak{M}^*

Proof. Only need to check forth property \square

Theorem 4.45. **S4** is strongly complete w.r.t. the class of partially ordered reflexive and transitive trees

Proof. If Σ is an **S4**-consistent set of formulas, and Σ^+ is an **S4**-MCS extending Σ , then $\mathfrak{M}^{\mathbf{S4}}, \Sigma^+ \Vdash \Sigma$. Moreover, as the **S4** axioms are canonical, $\mathfrak{M}^{\mathbf{S4}}$ is a reflexive transitive model. We now transform this model into the required partial order in two steps

Step 1. Let \mathfrak{M}^S be the submodel of $\mathfrak{M}^{\mathbf{S4}}$ generated by Σ^+ . Clearly this is a reflexive, transitive, point-generated model

Step 2. Let $\mathfrak{M}^* = (\vec{W}, R^*, \vec{V})$ be the reflexive transitive closure of the unraveling of \mathfrak{M}^S around Σ^+

By Lemma 4.44 \mathfrak{M}^S is a bounded morphic image of \mathfrak{M}^* under f , hence for all sequences $\vec{s} \in f^{-1}[\Sigma]$, we have $\mathfrak{M}^*, \vec{s} \Vdash \Sigma$. \square

K_t4.3 the tense logic generated by 4, .3_t and .3_r.

Definition 4.46. Let (T, R) be a transitive frame. A **cluster** on (T, R) is a subset C of T that is a maximal equivalent relation under R . A cluster is **simple** if it consists of a single reflexive point and **proper** otherwise.

Theorem 4.47. **K_t4.3** is strongly complete w.r.t. the class of strict total orders.

Proof. Let Σ be a $\mathbf{K}_t\mathbf{4.3}$ -consistent set of formulas; expand it to a $\mathbf{K}_t\mathbf{4.3}$ -MCS Σ . Let $\mathfrak{M} = (T, R, V)$ be a canonical model of $\mathbf{K}_t\mathbf{4.3}$. By the canonicity of the axioms, \mathfrak{M} is transitive and non-branching. Let $\mathfrak{M}^S = (S, R^S, V^S)$ be the submodel of \mathfrak{M} generated by Σ^+ ; \mathfrak{M}^S is a transitive and trichotomous model s.t. $\mathfrak{M}^S, \Sigma^+ \Vdash \Sigma$. But \mathfrak{M}^S may contain clusters, which we will bulldoze away

Step 1. Index the clusters in \mathfrak{M}^S by some suitable set I

Step 2. Define an arbitrary strict order $<^I$ on each cluster C_i

Step 3. Define C_i^\flat to be $C_i \times \mathbb{Z}$

Step 4. Define B , the set underlying the bulldozed model, to be

$$S^- \cup \bigcup_{i \in I} C_i^\flat$$

where S^- is the set $(S \setminus \bigcup_{i \in I} C_i)$ of points *not* belonging to any clusters

Step 5. Define a mapping $\beta : B \rightarrow S$ by $\beta(b) = b$ if $b \in S^-$; and $\beta(b) = s$ if $b = (s, z)$

Step 6. Define an ordering $<^b$ on B by $b <^b b'$ iff

- **either** $(b \in S^- \text{ or } b' \in S^-)$ and $\beta(b)R^S\beta(b')$
- **or** $b = (s, z)$ and $b' = (s', z')$ and
 - **either** s and s' belong to distinct clusters and $\beta(b)R^S\beta(b')$
 - **or** s and s' belong to the same cluster and $z <_{\mathbb{Z}} z'$
 - **or** s and s' belong to the same cluster C_i and $z = z'$ and $s <^i s'$

Step 7. Define a valuation V^b on $(B, <^b)$ by $b \in V(p)$ iff $\beta(b) \in V^S(p)$

Step 8. Define \mathfrak{M}^B , the **bulldozed model**, to be $(B, <^b, V^b)$ □

Claim: The mapping β is a surjective bounded morphism from $(B, <^b)$ to (S, R^S) and the model \mathfrak{M}^S is a bounded morphic image of \mathfrak{M}^B under β

Proof. □

Claim: $(B, <^b)$ is a strict total order

4.6 Step by Step

In what follows, consistency means $\mathbf{K}_t\mathbf{Q}$ -consistency and $\mathfrak{M}^c = (T^c, R^c, V^c)$ is this logic's canonical model. Furthermore we fix a maximal consistent set Σ ; the goal of our proof is to construct a model $\mathfrak{M} = (T, <, V)$ for Σ s.t. $(T, <)$ is an ordering isomorphic to $(\mathbb{Q}, <)$

Definition 4.48. A **network** is a triple $\mathcal{N} = (N, <, \nu)$ s.t. R is a binary relation on the set N , and ν is a labeling function mapping each point in N to a maximal consistent set

Definition 4.49. A network $\mathcal{N} = (N, <, \nu)$ is **coherent** if it satisfies:

1. $<$ is a strict total ordering
2. $\nu(s)R^c\nu(t)$ for all $s, t \in N$ s.t. $s < t$

A **network for** Σ is a network s.t. Σ is the label set of some node

2 is equivalent to the requirement that if $s < t$ then $F\phi \in \nu(s)$ for all $\phi \in \nu(t)$ and $P\phi \in \nu(t)$ for all $\phi \in \nu(s)$

Definition 4.50. A network $\mathcal{N} = (N, <, \nu)$ is **saturated** if it satisfies

1. $<$ is unbounded to the left and to the right
2. $<$ is dense
3. \mathcal{N} is modally saturated. That is, we demand that (F) if $F\psi \in \nu(s)$ for some $s \in N$, then there is some $t \in N$ s.t. $s < t$ and $\psi \in \nu(t)$ and (P) if $P\psi \in \nu(s)$ for some $s \in N$, then there is some $t \in N$ s.t. $t < s$ and $\psi \in \nu(t)$

A network is **perfect** if its both coherent and saturated

Definition 4.51. Let $\mathcal{N} = (N, <, \nu)$ be a network. The frame $\mathfrak{F}_{\mathcal{N}} = (N, <)$ is called the **underlying frame** of \mathcal{N} . The **induced valuation** $V_{\mathcal{N}}$ on \mathfrak{F} is defined by $V_{\mathcal{N}}(p) = \{s \in N \mid p \in \nu(s)\}$. The structure $\mathfrak{J}_{\mathcal{N}} = (\mathfrak{F}_{\mathcal{N}}, V_{\mathcal{N}})$ is the **induced model**

Lemma 4.52 (Truth Lemma). *Let \mathcal{N} be a countably infinite perfect network. Then for all formulas ψ , and all nodes s in N ,*

$$\mathfrak{J}_{\mathcal{N}}, s \Vdash \psi \quad \text{iff} \quad \psi \in \nu(s)$$

Moreover, $\mathfrak{F}_{\mathcal{N}}$ is isomorphic to the ordering of the rational numbers

Proof. Base case is from definition. Boolean case is from MCS. The coherency of \mathcal{N} drives the left to right implication, and saturation takes care of the other direction

The underlying frame of a perfect network must be a dense, unbounded, strict total ordering. Hence if it's countably infinite, it must be isomorphic to $(\mathbb{Q}, <)$ by Cantor's Theorem \square

It follows from Lemma 4.52 that we have reduced the task of finding a model for our MCS Σ to the quest for a countable, perfect network for Σ .

Definition 4.53. Let $\mathcal{N} = (N, <, \nu)$ be a network. An **S1-defect** of \mathcal{N} consists of a node $s \in N$ that has no successor, or no predecessor; an **S2-defect** is a pair (s, t) of nodes for which there is no intermediate point. An **S3-defect** consists of (F) a node s and a formula $F\psi \in \nu(s)$ for which there is no t in N s.t. $s < t$ and $\psi \in \nu(t)$ or (P) a node s and a formula $P\psi \in \nu(s)$ for which there is no t in N s.t. $t < s$ and $\psi \in \nu(t)$

Definition 4.54. Let $\mathcal{N}_0 = (N_0, <_0, \nu_0)$ and $\mathcal{N}_1 = (N_1, <_1, \nu_1)$ be two networks. We say that \mathcal{N}_1 **extends** \mathcal{N}_0 (notation $\mathcal{N}_1 \triangleright \mathcal{N}_0$) if $\mathfrak{F}_{\mathcal{N}_0}$ is a subframe of $\mathfrak{F}_{\mathcal{N}_1}$ and ν_0 agrees with ν_1 on N_0

Lemma 4.55 (Repair Lemma). *For any defect of a finite coherent network \mathcal{N} there is a finite coherent $\mathcal{N}' \triangleright \mathcal{N}$ lacking this defect*

Proof. Let $\mathcal{N} = (N, <, \nu)$ be a finite, coherent network and assume that \mathcal{N} has some defect.

S1-defects

S2-defects

Assume that there are nodes s and t in N for which there is no intermediate point. By coherence of \mathcal{N} , $\nu(s)R^c\nu(t)$, and by canonicity of the density axiom that there is some MCS Γ s.t. $\nu(s)R^c\Gamma R^c\nu(t)$. Hence take some *new* node u and define $\mathcal{N}' = (N', <', \nu')$ by

$$\begin{aligned} N' &:= N \cup \{u\} \\ <' &:= < \cup \{(x, u) \mid x \leq s\} \cup \{(u, x) \mid t \leq x\} \\ \nu' &:= \nu \cup \{(u, \Gamma)\} \end{aligned}$$

Now we need to prove \mathcal{N}' is coherent.

S3-defects

We only treat the P-defects. Assume that there is a node s in N and a formula $P\psi \in \nu(s)$ for which is no t in N s.t. $t < s$ and $\psi \in \nu(t)$

Let m be the unique point in N s.t.

1. $(m, P\psi)$ is an S3-defect in \mathcal{N}
2. for all $w < m$, $(w, P\psi)$ is *not* a defect.

Such an m must exist (it is either s itself or one of the finitely many points preceding s) and, we can repair $(m, P\psi)$ without problems by simply inserting the new point s' immediately before m

Choose some new point s' and let Ψ be an MCS containing ψ s.t. $\Psi R^c \nu(m)$; such a Ψ exists by the Existence Lemma for normal logics. Define $\mathcal{N}' = (N', <', \nu')$ as follows

$$\begin{aligned} N' &:= N \cup \{s'\} \\ <' &:= < \cup \{(x, s') \mid x < m\} \cup \{(s', x) \mid m \leq x\} \\ \nu' &:= \nu \cup \{(s', \Psi)\} \end{aligned}$$

It only remains to ensure that \mathcal{N}' satisfies the second coherency condition

Suppose $x < s'$. By construction $\nu(s') = \Psi R^c \nu(m)$, and by coherency of \mathcal{N} , $\nu(x) R^c \nu(m)$. But R^c is the canonical relation for $\mathbf{K}_t\mathbf{Q}$ - a relation with no branching to the left - hence either $\Psi R^c \nu(x)$, $\Psi = \nu(x)$ or $\nu(x) R^c \Psi$. We claim that the first two options are impossible. For if $\Psi R^c \nu(x)$ then $\psi \in \Psi$ would imply that $P\psi \in \nu(x)$ and this contradicts the minimality of m \square

Theorem 4.56. $\mathbf{K}_t\mathbf{Q}$ is strongly complete w.r.t. $(\mathbb{Q}, <)$

Proof. Choose some set $S = \{s_i \mid i \in \omega\}$ and enumerate the set of potential defects. \square

4.7 Rules for the Undefinable

Irreflexivity, although not definable in basic modal languages, **can** be characterized in an alternative sense:

If a temporal formula ψ is satisfiable on an irreflexive frame, then for any proposition letter p not occurring in ψ , the conjunction $(\neg Pp \wedge \neg Fp) \wedge \psi$ is also satisfiable on that frame

For if $\mathfrak{F}, V, s \models \psi$, then $\mathfrak{F}, V', s \models (\neg Pp \wedge p \wedge \neg Fp) \wedge \psi$ where V' is just like V except that it assigns the singleton $\{s\}$ to p .

Now by taking the contrapositive of the above statement, we turn it into a proof rule

(IRR) if $\vdash (\neg Pp \wedge p \wedge \neg Fp) \rightarrow \phi$ then $\vdash \phi$, provided p doesn't occur in ϕ

Note that on the class of strict total orders the formula $(\neg P\phi \wedge \phi \wedge \neg F\phi)$ is true at some state s iff s is the **only** state where ϕ holds (we need trichotomy and transitivity to ensure this). Call this formula $name(\phi)$

Definition 4.57. The logic $\mathbf{K}_t\mathbf{Q}^+$ is obtained by adding to $\mathbf{K}_t\mathbf{Q}$ the irreflexivity rule IRR. In what follows, consistency means $\mathbf{K}_t\mathbf{Q}^+$ -consistency, $\vdash \phi$ means that ϕ is provable in $\mathbf{K}_t\mathbf{Q}^+$, and so on. The canonical model for $\mathbf{K}_t\mathbf{Q}^+$ is denoted by \mathfrak{M}^c , the canonical relation by R^c

Definition 4.58. A maximal consistent set is called **witnessing** if it contains a formula of the form $name(\phi)$

Definition 4.59. Let W be a set of maximal consistent sets of formulas. Define $\mathfrak{M}^c|_W$ to be the submodel of the canonical model induced by W ; that is, $\mathfrak{M}^c|_W = (W, R, V)$ where R is the relation R^c restricted to W , and V is the canonical relation restricted to W

Definition 4.60. A set W of maximal consistent sets is called **diamond saturated** if it satisfies the requirement that for each $\Sigma \in W$ and each formula $F\psi \in \Gamma$ there is a set $\Psi \in W$ s.t. $\Sigma R^c \Psi$ and $\psi \in \Psi$, and the analogous condition holds for past formulas

Definition 4.61 (Truth Lemma). Let W be a diamond saturated set of maximal consistent sets of formulas. Then for any $\Gamma \in W$ and any formula ϕ :

$$\mathfrak{M}^c|_W, \Gamma \Vdash \phi \quad \text{iff} \quad \phi \in \Gamma$$

Proposition 4.62. Let ξ be some consistent formula. Then there is a countable, diamond saturated collection W of witnessing MCSs s.t. $\xi \in \Xi$ for some $\Xi \in W$

All this is done to ensure that in the limit we are dealing with a labeled graph meeting the requirements that

1. the label set of each point is an MCS
2. each label set contains a witness
3. if a formula of the form $F\phi$ ($P\phi$) belongs to the label set of some node, then there is an edge connecting this node to another one containing ϕ in its label set.

Finally W is defined as the range of this infinite labeling function

Approximations to W will be called **networks**: a network is a quadruple $\mathcal{N} = (N, E, d, \Lambda)$ s.t. (N, E) is a finite, undirected, connected and acyclic graph; d is a direction function mapping each edge (s, t) of the graph to either R or its converse R^\sim ; and Λ is a label function mapping each node of N to a *finite* set of formulas

We first want to formulate coherence conditions on networks and define the notion of a defect of network w.r.t. its ideal, W . Since we are working in the basic temporal similarity type - that is, we have diamonds both for looking along R and along R^\sim - there is an obvious way of describing the network, from each of its nodes. Let $\mathcal{N} = (N, E, d, \Lambda)$ be some network,

and let s and t be two adjacent nodes of \mathcal{N} . We use the following notational conventions

$$\langle st \rangle := \begin{cases} F & \text{if } d(s, t) = R \\ P & \text{if } d(t, s) = R^\sim \end{cases}$$

Moreover, let $E(s)$ denote the set of nodes adjacent to s . Finally, we let $\lambda(s)$ denote the conjunction $\bigwedge \Lambda(s)$. Define

$$\begin{aligned} \Delta(\mathcal{N}, s) &:= \lambda(s) \wedge \bigwedge_{v \in E(s)} \langle sv \rangle \theta(\mathcal{N}, v, s) \\ \theta(\mathcal{N}, t, s) &:= \lambda(t) \wedge \bigwedge_{s \neq v \in E(t)} \langle tv \rangle \theta(\mathcal{N}, v, t) \end{aligned}$$

In other words, $\Delta(\mathcal{N}, s)$ starts with a local description $\lambda(s)$ of s and then proceeds to its neighbors. For each neighbor v , $\Delta(\mathcal{N}, s)$ writes a function operator if $d(s, v) = R$ and then starts to describe the network after v by calling θ . $\theta(\mathcal{N}, v, s)$ first gives a local description $\lambda(v)$ of v , and then recursively proceeds to the neighbors of v - except for s . The omission of s , together with the finiteness and acyclicity of the graph, ensures that we end up with a finite formula

Lemma 4.63. *For any network \mathcal{N} and any two nodes s, t in \mathcal{N} , $\Delta(\mathcal{N}, s)$ is consistent iff $\Delta(\mathcal{N}, t)$ is consistent*

Proof. By the connectedness of \mathcal{N} it is sufficient to prove the Lemma for adjacent s and t .

So suppose that s and t are adjacent; without loss of generality assume that $d(s, t) = R$. Since \mathcal{N} is fixed it won't lead to confusion if we abbreviate $\Delta(\mathcal{N}, x)$ by $\Delta(x)$ and $\theta(\mathcal{N}, x, y)$ by $\theta(x, y)$. Then by definition, $\Delta(s)$ is given by

$$\begin{aligned} \Delta(s) &= \lambda(s) \wedge \bigwedge_{u \in E(s)} \langle su \rangle \theta(u, s) \\ &= \lambda(s) \wedge F\theta(t, s) \wedge \bigwedge_{t \neq u \in E(s)} \langle su \rangle \theta(u, s) \\ &= F\theta(t, s) \wedge \theta(s, t) \end{aligned}$$

Likewise we can show that

$$\Delta(t) = \theta(t, s) \wedge P\theta(s, t)$$

But it is a general property of any logic extending \mathbf{K}_t that for any two formulas α and β , $F\alpha \wedge \beta$ is consistent iff $\alpha \wedge P\beta$ is consistent \square

We call a network \mathcal{N} **coherent** if $\Delta(\mathcal{N}, s)$ is consistent for each of its nodes.

A **defect** of a network is either

1. a pair (s, ϕ) s.t. neither ϕ nor $\neg\phi$ belongs to $\Lambda(s)$
2. a pair $(s, F\phi)$ s.t. $F\phi \in \Lambda(s)$ while there is no witness for this in the sense that $\phi \in \Lambda(t)$ for some node t with Est and $d(s, t) = R$
3. a similar pair $(s, P\phi)$
4. a node s without a name; that is, $name(\phi) \in \Lambda(s)$ for no formula

A network \mathcal{N}' **extends** a network \mathcal{N} ($\mathcal{N}' \triangleright \mathcal{N}$) if NN' while $E = E' \cap N \times N$, $d = d'|_N$ and $\Lambda(s) \subseteq \Lambda'(s)$ for each node s

Lemma 4.64. *For any defect of a finite coherent network \mathcal{N} there is a finite, coherent $\mathcal{N}' \triangleright \mathcal{N}$ lacking this defect*

Proof. *D4-defects*

If $\Delta(\mathcal{N}, s)$ is consistent, then there is a propositional variable p that doesn't occur in any of the label sets. And we use the IRR-rule to show that the formula $\Delta(\mathcal{N}, s) \wedge name(p)$ is consistent \square

Now we return to the proof of Proposition 4.62

Proof. By a standard step-by-step construction we can define a sequence $(\mathcal{N}_i)_{i \in \mathbb{N}}$ of networks s.t.

1. \mathcal{N}_0 is a one-node network with label set $\{\xi\}$
2. \mathcal{N}_j extends \mathcal{N}_i whenever $i < j$
3. for every defect of any network \mathcal{N}_i there is a network \mathcal{N}_j with $j > i$ lacking this defect

Let N be the set $\bigcup_{i \in \mathbb{N}} N_i$; and for $s \in N$, define $\Lambda(s) = \bigcup_{i \in \mathbb{N}} \Lambda_i(s)$. We claim that for every $s \in N$, $\Lambda(s)$ is a witnessing MCS. Assume $\Lambda(s)$ is not consistent; then there are formulas ϕ_1, \dots, ϕ_n in $\Lambda(s)$ s.t. $\phi_1 \wedge \dots \wedge \phi_n$ is inconsistent. By construction, there must be a $k \in \mathbb{N}$ s.t. each ϕ_i belongs to $\Lambda_k(s)$. But this contradicts the consistency of $\Delta(\mathcal{N}_k, s)$ and hence the coherency of \mathcal{N}_k

Finally define W as the range of Λ . The preceding paragraphs show that W is a collection of witnessing MCSs. By our definition of \mathcal{N}_0 , it follows that ξ belongs to some MCS in W

Now let $F\phi$ be some formula in $\Gamma \in W$. By definition, there is some $s \in N$ s.t. $\Gamma = \Lambda(s)$, and thus, some $i \in \mathbb{N}$ s.t. $F\phi \in \Lambda_i(s)$. By our construction there is some $j \geq i$ and some $t \in N_j$ s.t. E_jst and $\phi \in \Lambda_j(t)$. It follows that $\phi \in \Lambda(t)$, so it remains to prove that $\Lambda(s)R^c\Lambda(t)$. Suppose otherwise, then there is a formula $\psi \in \Lambda(t)$ s.t. $F\psi \notin \Lambda(s)$. Since $\Lambda(s)$ is an MCS, this implies $\neg F\psi \in \Lambda(s)$. Now let $k \in \mathbb{N}$ be large enough that $\psi \in \Lambda_k(t)$ and $\neg F\psi \in \Lambda_k(s)$. From this it is immediate that $\Delta(\mathcal{N}_k, s)$ is inconsistent \square

Lemma 4.65. *Let W be a diamond saturated collection of witnessing maximal consistent sets of formulas, and let $<$ denote the canonical relation R^c restricted to W . Then the frame $(W, <)$ is a non-branching, unbounded, dense strict ordering*

Proof. Let W and $<$ be as in the frame. $(W, <)$ is a subframe of the canonical frame; hence it inherits every **universal** property of \mathfrak{T} , such as transitivity and non-branching. Irreflexivity follows from the fact that $\Gamma R^c \Gamma$ for no witnessing Γ .

Unboundedness is not a universal condition. As $F\top$ and $P\top$ are theorems of the logic and hence belong to every maximal consistent set

For density. Assume that Γ and Δ are two MCSs s.t. $\Gamma < \Delta$. We have to find an MCS Θ in W that lies between Γ and Δ . Let δ be the formula s.t. $\text{name}(\delta) \in \Delta$. It follows from $\Gamma < \Delta$ that $F\text{name}(\delta) \in \Gamma$, so using the density axiom, we find that $FF\text{name}(\delta) \in \Gamma$. From this we may infer the existence of an MCS $\Theta \in W$ with $\Gamma < \Theta$ and $F\text{name}(\delta) \in \Theta$.

But is $\Theta < \Delta$. Note that since $<$ is non-branching to the right, we already know that $\Theta < \Delta$ or $\Theta = \Delta$ or $\Delta < \Theta$. But its clear that $F\delta \in \Theta$ and $\neg F\delta \in \Delta$. Neither is it possible that $\Delta < \Theta$ for suppose otherwise. It would follow that $F\delta \in \Delta$ that $FF\delta \in \Delta$, so by transitivity axiom, $F\delta \in \Delta$ \square

Theorem 4.66. $K_t Q^+$ is complete w.r.t. $(\mathbb{Q}, <)$

Proof. Given any consistent formula ξ , construct a countable, diamond saturated set W of witnessing MCSs for ξ . Then by the Truth Lemma, ξ is satisfiable at some MCS Ξ in the model $\mathfrak{M}^C|_W$ induced by W ; and by Lemma 4.65 this model is based on a non-branching, unbounded, dense, strict ordering. But then the subframe generated by Ξ is based on a countable, dense, unbounded, strict total order and hence is isomorphic to the ordering of the rationals \square

4.8 Finitary Methods I

A frame for this language is a transition system $\mathfrak{F} = (W, R_\pi)_{\pi \in \Pi}$, but we only interested in **regular frames**, that is, frames s.t. for all programs π, π_1

and π_2 :

$$\begin{aligned} R_{\pi_1 \cup \pi_2} &= R_{\pi_1} \cup R_{\pi_2} \\ R_{\pi_1; \pi_2} &= R_{\pi_1}; R_{\pi_2} \\ R_{\pi^*} &= (R_{\pi})^* \end{aligned}$$

A formula ϕ is a PDL-validity (written $\models \phi$) if its valid on all regular frame

The collection of PDL-validities is not compact; consider the set

$$\Sigma = \{ \langle a^* \rangle p, \neg p, \neg \langle \rangle p, \neg \langle a \rangle \langle a \rangle p, \dots \}$$

The compactness failure indicates that a **strong** completeness result will be out of reach

Definition 4.67. A logic Λ in the language of propositional dynamic logic is a **normal propositional dynamic logic** if it contains every instance of the following axiom schemas

1. $[\pi](p \rightarrow q) \rightarrow ([\pi]p \rightarrow [\pi]q)$
2. $\langle \pi \rangle p \leftrightarrow \neg[\pi]\neg p$
3. $\langle \pi_1; \pi_2 \rangle p \leftrightarrow \langle \pi_1 \rangle \langle \pi_2 \rangle p$
4. $\langle \pi_1 \cup \pi_2 \rangle p \leftrightarrow \langle \pi_1 \rangle p \vee \langle \pi_2 \rangle p$
5. $\langle \pi^* \rangle p \leftrightarrow (p \vee \langle \pi \rangle \langle \pi^* \rangle p)$
6. $[\pi^*](p \rightarrow [\pi]p) \rightarrow (p \rightarrow [\pi^*]p)$

and is closed under modus ponens, generalization and uniform substitution. We call the smallest normal propositional dynamic logic **PDL**. In this section $\vdash \phi$ means ϕ is a theorem of **PDL**

axiom 6 is **induction** axiom.

Exercise 4.8.1. Prove induction axiom is not canonical

Proof. Let $\psi = [\pi^*](p \rightarrow [\pi]p) \rightarrow (p \rightarrow [\pi^*]p)$. □

Definition 4.68 (Fischer-Ladner Closure). Let X be a set of formulas. Then X is **Fischer-Ladner closed** if it is closed under subformulas and satisfied the following additional constraints

1. If $\langle \pi_1; \pi_2 \rangle \phi \in X$ then $\langle \pi_1 \rangle \langle \pi_2 \rangle \phi \in X$
2. If $\langle \pi_1 \cup \pi_2 \rangle \phi \in X$ then $\langle \pi_1 \rangle \phi \vee \langle \pi_2 \rangle \phi \in X$
3. If $\langle \pi^* \rangle \phi \in X$ then $\langle \pi \rangle \langle \pi^* \rangle \phi \in X$

If Σ is any set of formulas then $\text{FL}(\Sigma)$ is the smallest set of formulas containing Σ that is Fischer-Ladner closed

Given a formula ϕ , we define

$$\sim\phi = \begin{cases} \psi & \text{if } \phi \text{ is of the form } \neg\psi \\ \neg\phi & \text{otherwise} \end{cases}$$

A set of formulas X is **closed under single negations** if $\sim\phi$ belongs to X whenever $\phi \in X$

We define $\neg\text{FL}(\Sigma)$ the **closure of Σ** , as the smallest set containing Σ which is Fischer-Ladner closed and closed under single negations

Definition 4.69 (Atoms). Let Σ be a set of formulas. A set of formulas A is an **atom** over Σ if it is maximal consistent subset of $\neg\text{FL}(\Sigma)$. $\text{At}(\Sigma)$ is the set of all atoms over Σ

Lemma 4.70. Let \mathcal{M} be the set of all MCSs, and Σ any set of formulas. Then

$$\text{At}(\Sigma) = \{\Gamma \cap \neg\text{FL}(\Sigma) \mid \Gamma \in \mathcal{M}\}$$

Proof. For any atom $A \in \text{At}(\Sigma)$, it can be extended to MCS. We need to prove $A \subseteq \neg\text{FL}(\Sigma)$. This can be done by the induction on the complexity of any formula ϕ . \square

Lemma 4.71. If $\phi \in \neg\text{FL}(\Sigma)$ and ϕ is consistent, then there is an $A \in \text{At}(\Sigma)$ s.t. $\phi \in A$

Proof. If Σ is infinite, the result is exactly Lindenbaum's Lemma. Otherwise, let $\widehat{A} = \bigwedge_{\phi \in A} \phi$

Enumerate the elements of $\neg\text{FL}(\Sigma)$ as $\sigma_1, \dots, \sigma_m$. Let A_0 be $\{\phi\}$. Suppose that A_n has been defined, where $n < m$. We have that

$$\vdash \widehat{A}_n \leftrightarrow (\widehat{A}_n \wedge \sigma_{n+1}) \vee (\widehat{A}_n \wedge \sim\sigma_{n+1})$$

as this a propositional tautology, thus either $A_n \cup \{\sigma_{n+1}\}$ or $A_n \cup \{\sim\sigma_{n+1}\}$ is consistent. Let A_{n+1} be the consistent extension, and let A be A_m . Then A is an atom containing ϕ \square

Definition 4.72 (Canonical Model over Σ). Let Σ be a finite set of formulas. The **canonical model over** Σ is the triple $(At(\Sigma), \{S_\pi^\Sigma\}_{\pi \in \Pi}, V^\Sigma)$ where for all propositional variables p , $V^\Sigma(p) = \{A \in At(\Sigma) \mid p \in A\}$, and for all atoms $A, B \in At(\Sigma)$ and all programs π ,

$$AS_\pi^\Sigma B \text{ if } \widehat{A} \wedge \langle \pi \rangle \widehat{B} \text{ is consistent}$$

V^Σ is called the **canonical valuation**, and the S_π are called the **canonical relations**.

Definition 4.73 (Regular PDL-model over Σ). Let Σ be a finite set of formulas. For all basic programs a , define R_a^Σ to be S_a^Σ . For the complex programs, inductively define the PDL-relations R_π^Σ in the usual way using unions, compositions and reflexive transitive closures. Finally define \mathfrak{M} the **regular PDL-model over** Σ , to be $(At(\Sigma), \{R_\pi^\Sigma\}_{\pi \in \Pi}, V^\Sigma)$ where V^Σ is the canonical valuation.

Lemma 4.74 (Existence Lemma for Basic Programs). *Let A be an atom, and let a be a basic program. Then for all formulas $\langle a \rangle \psi$ in $\neg FL(\Sigma)$, $\langle a \rangle \psi \in A$ iff there is a $B \in At(\Sigma)$ s.t. $AR_a B$ and $\psi \in B$*

Proof. Right to the left. Suppose there is a $B \in At(\Sigma)$ s.t. $AR_a B$ and $\psi \in B$. As R_a and S_a are identical for basic programs, we have that $AS_a B$, thus $\widehat{A} \wedge \langle a \rangle \widehat{B}$ is consistent. As ψ is one of the conjuncts in \widehat{B} , $\widehat{A} \wedge \langle a \rangle \psi$ is consistent. As $\langle a \rangle \psi$ is in $\neg FL(\Sigma)$ it must also be in A , for A is an atom and hence maximal consistent in $\neg FL(\Sigma)$

Left to right. Suppose $\langle a \rangle \psi \in A$. Enumerate the formulas in $\neg FL(\Sigma)$ as $\sigma_1, \dots, \sigma_2$. Define B_0 to be $\{\psi\}$. Suppose as an inductive hypothesis that B_n is defined s.t. $\widehat{A} \cap \langle a \rangle \widehat{B}_n$ is consistent. We have

$$\vdash \langle a \rangle \widehat{B}_n \leftrightarrow \langle a \rangle ((\widehat{B}_n \wedge \sigma_{n+1}) \vee (\widehat{B}_n \wedge \sim \sigma_{n+1}))$$

thus

$$\vdash \langle a \rangle \widehat{B}_n \leftrightarrow (\langle a \rangle (\widehat{B}_n \wedge \sigma_{n+1}) \vee \langle a \rangle (\widehat{B}_n \wedge \sim \sigma_{n+1}))$$

Therefore either for $B' = B_n \cup \{\sigma_{n+1}\}$ or for $B' = B_n \cup \{\sim \sigma_{n+1}\}$ we have that $\widehat{A} \wedge \langle a \rangle \widehat{B}'$ □

Lemma 4.75. *For all programs π , $S_{\pi^*} \subseteq (S_\pi)^*$*

Proof. We need to show that for all programs π , if $AS_{\pi^*} B$ then there is a finite sequence of atoms C_0, \dots, C_n s.t. $A = C_0 S_\pi C_1, \dots, C_{n-1} S_\pi C_n = B$. Let \mathcal{D} be the set of all atoms reachable from A by such a sequence. We will show that $B \in \mathcal{D}$ □