

# Lecture notes

## Einführung in die Logik 2024W

This is a summary of the material discussed in the lecture "Mathematische Logik". It is still a work in progress and there **may be mistakes** in this work. If you find any, feel free to let me know and I will correct them

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## List of Abbreviations

prop.	-	propositional . . . . .	2
exp.	-	expression(s) . . . . .	2
sent.	-	sentence(s) . . . . .	3
seq.	-	sequence . . . . .	3
TA	-	truth assignment . . . . .	3
fla.	-	formula . . . . .	3
TV	-	truth value . . . . .	3
taut.	-	tautological . . . . .	4
w/	-	with . . . . .	4
lp / rp	-	left / right parenthesis . . . . .	5
i.e.	-	id est (that is) . . . . .	9
SA	-	substitutable . . . . .	20

## CHAPTER 1

## Propositional logic

Language **Definition 1.1. Language of PL:** The Language of Propositional logic is a set containing

- logical symbols: consisting of the **sentential connective** symbols  $\neg, \wedge, \vee, \rightarrow, \leftrightarrow$  and parenthesis  $(, )$
- non-logical symbols:  $A_1, A_2, A_3, \dots$  (also called sentential atoms, variables)

from which we assume (for unique readability) that no symbol is a finite sequence of any other symbols.

**Note:**

1. The role of the logical symbols doesn't change, the sentential atoms we see as variables, they function as placeholders or variables.
2. we assumed the set of non-logical symbols is countable, for most of our conclusions you could use any set of prop. atoms of any size

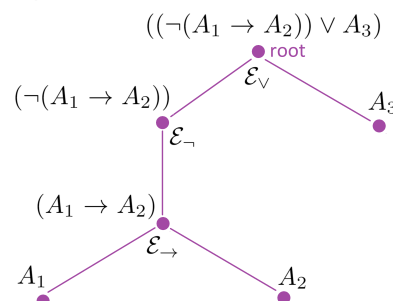
expression **Definition 1.2. Expression / prop. sentence:** An **expression** is a any finite sequence of symbols We define **grammatically correct exp.** recursive

1. every prop. atom is a prop. sentence
2. if  $\alpha, \beta$  are prop. sentences, then also  $\neg\alpha, \alpha \wedge \beta, \alpha \vee \beta, \alpha \rightarrow \beta, \alpha \leftrightarrow \beta$
3. nothing else (in particular  $\emptyset$  is not a prop. fla.)

prop. fla. and call them **prop. sentences** or **prop. fla.** Equivalently stated every prop. sentence is built up by applying finitly many formula building operations on atoms and the prop. sent. returned from building operations.

$$\mathcal{E}_{\neg}, \mathcal{E}_{\neg}(\alpha) := (\neg\alpha) \text{ for any prop. fla. } \alpha \text{ and similarly for } \mathcal{E}_{\wedge}, \mathcal{E}_{\vee}, \mathcal{E}_{\rightarrow}, \mathcal{E}_{\leftrightarrow}$$

This allows us to symbolize the **expression tree** (Here for example for  $((\neg(A_1 \rightarrow A_2)) \vee A_3)$ )



We will return to these construction trees in 1.2, where we answer the question of what truth value a given prop. sentence might have.

**Definition 1.3. Construction sequence:** Given a prop. sentence  $\alpha$  a **construction sequence** of  $\alpha$  is a finite sequence  $\langle \alpha_1, \dots, \alpha_{n-1}, \alpha \rangle$  such that for all  $i \leq n$  the following holds

construction sequence

- $\alpha_i$  is a sentential atom
- or  $\alpha_i = \mathcal{E}_{\neg}(\alpha_j)$  for some  $j < i$
- or  $\alpha_i = \mathcal{E}_{\Box}(\alpha_j, \alpha_k)$  for some  $j, k < i$  and  $\Box \in \{\wedge, \vee, \rightarrow, \leftrightarrow\}$

**Definition 1.4. Closedness of a set:** Let  $S$  be a set. We say  $S$  is **closed** under an  $n$ -ary operational symbol  $f$  iff for all  $s_1, s_2, \dots, s_n \in S$  it holds  $f(s_1, s_2, \dots, s_n) \in S$

closure

**Induction principle:** Suppose  $S$  is a set of prop. sentences containing all prop. atoms and closed under the 5 formula building operations, then  $S$  is the set of all prop. sentences.

*Proof.* let  $PS$  = set of all prop. sent.

$S \subseteq PS$ : is clear

$S \supseteq PS$ : let  $\alpha \in PS$  then  $\alpha$  has a construction seq.  $\langle \alpha_1, \dots, \alpha_{n-1}, \alpha \rangle$  and  $\alpha_1 \in S$  lets assume that  $\alpha_i$  for  $i \leq k < n$  is in  $S$  then  $\alpha_{k+1}$  is either an atom and therefore in  $S$  or its obtained by one of the formula building operations from the and therefore  $\alpha_{k+1} \in S$

□

## 1.1 TRUTH ASSIGNMENTS

The interpretation of a prop. atom is either true or false, denoted by 0/1 or  $T/F$  or  $\top/\perp$ . A truth assignment is simply any map  $\nu : S \mapsto \{0, 1\}$ , where  $S$  is a map of propositional atoms. Our goal is going to be to extend any truth assignment  $\nu$  to a function  $\bar{\nu} : \bar{S} \mapsto \{0, 1\}$ , where  $\bar{S}$  is the closure of  $S$  under the 5 fla. building operations.

**Definition 1.5. Truth assignment:** Let  $\{0, 1\}$  be the set of truth values. A truth assignment (TA) for a set  $S$  of prop. atoms is a map  $\nu : S \rightarrow \{0, 1\}$

Truth assignment  
TA

We now want to extend  $\nu$  to  $\bar{\nu} : \bar{S} \rightarrow \{0, 1\}$ , where  $\bar{S}$  is the closure of  $S$  under the 5 fla. building operations such that for all propositional atoms  $A \in S$  and propositional formulas  $\alpha, \beta$  in  $\bar{S}$

1.  $\bar{\nu}(A) = \nu(A)$
2.  $\bar{\nu}(\neg\alpha) = 1 - \nu(\alpha)$
3.  $\bar{\nu}(\alpha \wedge \beta) = \begin{cases} 1 & \text{iff } \bar{\nu}(\alpha) = 1 = \bar{\nu}(\beta) \\ 0 & \text{otherwise} \end{cases}$
4.  $\bar{\nu}(\alpha \vee \beta) = \begin{cases} 1 & \text{iff } \bar{\nu}(\alpha) = 1 \text{ or } \bar{\nu}(\beta) = 1 \\ 0 & \text{otherwise} \end{cases}$
5.  $\bar{\nu}(\alpha \rightarrow \beta) = \begin{cases} 1 & \text{iff } \bar{\nu}(\alpha) = 0 \text{ or } \bar{\nu}(\beta) = 1 \\ 0 & \text{otherwise} \end{cases}$
6.  $\bar{\nu}(\alpha \leftrightarrow \beta) = \begin{cases} 1 & \text{iff } \bar{\nu}(\alpha) = \bar{\nu}(\beta) \\ 0 & \text{otherwise} \end{cases}$

We also want the extention to be unique, that is

**Theorem 1.1. Unique readability:** For all TA  $\nu$  for a set  $S \exists! \bar{\nu} : \bar{S} \rightarrow \{0, 1\}$  satisfying the above properties

We will proof this later

satisfy  
satisfiable

**Definition 1.6. Satisfaction:** A TA  $\nu$  satisfies a prop. sent.  $\alpha$  if  $\bar{\nu}(\alpha) = 1$  (that is, provided that every atom of  $\alpha$  is in the domain of  $\nu$ ). We call  $\alpha$  satisfiable if there exists a TA that satisfies it.

taut. implication  
 $\models$

**Definition 1.7. Tautological implication:** Let  $\Sigma$  be a set of prop. sent. and  $\alpha$  a prop. sent. then we say:  $\Sigma$  tautologically implies  $\alpha$  if for all TA that satisfy  $\Sigma$ ,  $\alpha$  is also satisfied and we write  $\Sigma \models \alpha$ . If  $\Sigma = \{\beta\}$ , we simply write  $\beta \models \alpha$ . If  $\Sigma = \emptyset$  then  $\alpha$  is called a **tautology** and we write  $\models \alpha$  instead of  $\emptyset \models \alpha$ .  
 $\alpha, \beta$  are called **tautologically equivalent** iff  $\alpha \models \beta$  and  $\beta \models \alpha$ , we then write  $\alpha \models \beta$

**Note:** In other words, tautological implication  $\Sigma \models \alpha$  means that you can not find a TA, that satisfy all members of  $\Sigma$  but not  $\alpha$ . A tautology is satisfied by every TA. Suppose there is no TA that satisfies  $\Sigma$ , then we have  $\Sigma \models \alpha$  for every prop. sent.  $\alpha$

**Example 1.1. :**  $\{\neg A \vee B\} \models A \rightarrow B$

**Note:** In order to check if a prop. sent. is satisfiable we need to check  $2^N$  TAs, where  $N = \#$  of atoms. It is unknown if this can be done by an algorithm in polynomial time. Answering this would settle the debate whether  $P = NP$

However we can find a way to reduce satisfiability of an infinite set  $\Sigma$  of prop. sent. to all finite subsets of  $\Sigma$ . There later will be a more elementary proof of the compactness theorem, this proof is not part of the exam.

**Theorem 1.2. Compactness theorem:** Let  $\Sigma$  be an infinite set of prop. sent. such that

$$\forall \Sigma_0 \subseteq \Sigma, \Sigma_0 \text{ finite} \exists \text{ TA satisfying every member of } \Sigma_0 \quad (\text{finite satisfiability})$$

then there is a TA satisfying every member of  $\Sigma$ .

*Proof.* using topology: We have our infinite set of prop. sent. which satisfies above condition. One way to look at TA is as a sequence of 0 and 1. Let  $\mathcal{A} = \{A_0, A_1, \dots\}$  be the set of all prop. atoms. We are going to identify the truth assignments on  $\mathcal{A}$  with elements in  $\{0, 1\}^{\mathcal{A}} := \{f : \mathcal{A} \rightarrow \{0, 1\}\}$  (the set of all TAs) This is a topological space with product topology, on which we will define the basic open sets (called cylinders) by:  $U \subseteq \{0, 1\}^{\mathcal{A}}$  is a cylinder, if it holds  $p_n(U) = \{0, 1\}$  for all but finite many  $n$ , where  $p_n$  is the  $n$ -th projections. This means  $U$  is a cylinder if the truth values of its elements are at finitely many places fixed, and are arbitrary on everything else.

**Note:** These basic open sets are also closed. We now define the open sets as unions of basic open sets. The idea is to use Tychonoff's Theorem which tells us that  $\{0, 1\}^{\mathcal{A}}$  is compact. i.e. the intersection of a family of closed subsets w/ the finite intersection property (FIP) is non-empty. Finite intersection property means the intersection of finitely many sets is non-empty.

For  $\alpha \in \Sigma$  let  $T_\alpha \subseteq \{0, 1\}^{\mathcal{A}}$  be the set of TA that satisfy  $\alpha$ . This  $T_\alpha$  is a finite union of cylinders, bc. it only depends on finitely many assignments, hence  $T_\alpha$  is closed. For the family  $\{T_\alpha : \alpha \in \Sigma\}$  of closed sets we have (FIP). Tychonoff tells us, that  $\bigcup_{\alpha \in \Sigma} T_\alpha \neq \emptyset$  so there is a TA satisfying  $\Sigma$ .  $\square$

For a list of tautologies: useful might be book p. 26-27

## 1.2 A PARSING ALGORITHM

To prove **Theorem 1.1** We essentially need to show that we have enough parenthesis to make the reading of a prop. sent. unique. That is given a TA  $v$  there is at most one truth value we can assign to a prop. sent.

**Lemma 1.1. :** Every prop. sent. has the same number of left and right parenthesis.

*Proof.* Let  $M$  = set of prop. sent. w/  $\#$  left parenthesis =  $\#$  right parenthesis and  $PS$  = set of all prop. sent. We have  $M \subseteq PS$ . Since atoms have no parenthesis, they are in  $M$ . we just need to show that  $M$  is closed under the 5 construction operations.  
 $\mathcal{E}_\neg = (\neg\alpha) \dots$   $\square$

**Lemma 1.2. :** No proper initial segment of a prop. sent. is itself a prop. sent.

*Proof.* Let  $\alpha = \alpha_1\alpha_2\ldots\alpha_n$  be a prop. sent. By proper initial segment we understand  $\beta = \alpha_1\ldots\alpha_i$  for  $1 \leq i < n$ . We will prove that every proper initial segment has an excess of left parenthesis, then we use the previous lemma. Let  $PS$  = set of all prop. sent. and  $PF$  = set of prop. sent. s.t. no proper initial segment has  $\#$  left parenthesis =  $\#$  right parenthesis, we will prove that these sets are the same.

Let  $\alpha \in PF$ . By induction over the fla. building operations

- Atoms: since the empty sequence is no prop. sent. they have no proper initial segment.
- If the above is true for  $\alpha, \beta$  then the proper initial segments of  $(\neg\alpha)$  are of the form

$(\neg\alpha$   
 $(\neg\alpha'$  where  $\alpha'$  is a proper initial segment of  $\alpha$   
 $($  or  
 $(\neg$

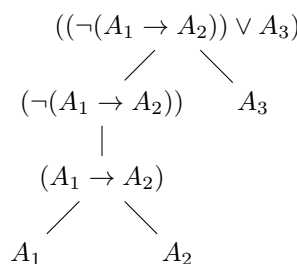
Therefore  $\mathcal{E}_\neg$  preserves this property and under  $\mathcal{E}_\wedge, \mathcal{E}_\vee, \mathcal{E}_\rightarrow, \mathcal{E}_{\leftrightarrow}$  this is also the case.  $\square$

### Parsing algorithm

We now give a parsing algorithm procedure. For input we take some expression  $\tau$  and the algorithm will determine if  $\tau$  is a prop. sent. If so, it will generate a unique construction tree (in form of a rooted tree) for  $\tau$ . (i.e. the construction tree gives us a unique readability) That there is a unique way to perform the algorithm is implied by [Lemma 1.2](#)

0. create the root and label it  $\tau$
1. HALT if all leaves are labeled w/ prop. atom and return: " $\tau$  is a prop. sent."
2. select a leaf of the graph which is not labeled w/ prop. atom
3. if the first symbol of label under consideration is not a left parenthesis, then halt and return: " $\tau$  is not a prop. sent."
4. if the second symbol of the label is " $\neg$ " then GOTO 6.
5. scan the expression from left to right  
if we reach a proper initial segment of the form " $(\beta$ " where  $\#lp(\beta) = \#rp(\beta)$  and  $\beta$  is followed by one of thesection  $\wedge, \vee, \rightarrow, \leftrightarrow$  and the remainder of the expression is of the form  $\beta'$ , where  $\#lp(\beta') = \#rp(\beta')$   
Then: create two child nodes (left,right) to the selected element and label them (left  $:= \beta$ , right  $:= \beta'$ ) GOTO 1.  
Else: HALT and return " $\tau$  is not a prop. sent."
6. if the expression is of the form  $(\neg\beta)$  where  $\#lp(\beta) = \#rp(\beta)$   
Then: construct one childnode and label it  $\beta$  and GOTO 1.  
Else: HALT and return: " $\tau$  is not a prop. sent."

**Example 1.2. :** The parsing algorithm applied to  $((\neg(A_1 \rightarrow A_2)) \vee A_3)$  returns the following construction tree.



## Correctness of the parsing algorithm

- The algorithm always halts, because the length of a child's label is less than the label of a parent.
- If the algorithm halts with the conclusion that  $\tau$  is a prop. sent. then we can prove inductively (starting from the leaves) that each label is a prop. sent
- Unique way to make choices in the algorithm: in particular  $\beta, \beta'$  in step 5. If there was a shorter choice for  $\beta$  it would be a proper initial segment of  $\beta$  but such prop. sent. can not exist. (This also works under the assumption that a longer choice exists).
- rejections are made correctly

Back to proving the existence and uniqueness of  $\bar{\nu}$  in [Theorem 1.1](#). Let  $\alpha$  be a prop. sent. of  $\bar{S}$ . We apply the parsing algorithm to  $\alpha$  to get a unique construction tree. For the leaves, use  $\nu$  to get the truth values then work our way up using the conditions (1-6) in [Definition 1.5](#).

## 1.3 INDUCTION AND RECURSION

### Generalization of induction principle:

Let  $U$  be a set and  $B \subseteq U$  our initial set.  $\mathcal{F} = \{f, g\}$  a class of functions containing just  $f$  and  $g$ , where

$$f : U \times U \rightarrow U, \quad g : U \rightarrow U$$

We want to construct the smallest subset  $C \subseteq U$  such that  $B \subseteq C$  and  $C$  is closed under all elements of  $\mathcal{F}$ .

**Definition 1.8. Closedness, Inductiveness:** We say  $\mathcal{S} \subseteq U$  is

- |           |  |
|-----------|--|
| closed    | • <b>closed</b> under $f$ and $g$ iff for all $x, y \in \mathcal{S}$ it holds $f(x, y) \in \mathcal{S}$ and $g(x) \in \mathcal{S}$ |
| inductive | • <b>inductive</b> if $B \subseteq \mathcal{S}$ and $\mathcal{S}$ is closed under $\mathcal{F}$                                    |

One way is from the top down

$$C^* := \bigcap_{B \subseteq S \text{ inductive}} S$$

Another is from bottom up: We call  $C_1 := B$ ,

$$C_i := C_{i-1} \cup \{f(x, y) : x, y \in C_{i-1}\} \cup \{g(x) : x \in C_{i-1}\}$$

and  $C_* := \bigcup_{n \geq 1} C_n$ . Exercise: show that  $C^* = C_* =: C$ .

**Example 1.3. :**

1. Let  $U$  be the set of all expressions,  $B$  the set of atoms and  $\mathcal{F} = \{\mathcal{E}_{\Box} : \Box \in \{\neg, \wedge, \vee, \rightarrow, \leftrightarrow\}\}$ . Then  $C$  would be the set of all propositional formulas.
2. Let  $U$  be  $\mathbb{R}$ ,  $B$  the set containing 0 and  $\mathcal{F} = \{S\}$ ,  $S(x) = x + 1$ . Then  $C$  would be the set of the natural numbers.

### Induction principle

$C$  generated from  $B$  by use of elements of  $\mathcal{F}$  if  $S \subseteq C$  such that  $B \subseteq S$  and  $S$  is closed under all elements of  $\mathcal{F}$ , then  $S = C$

*Proof.*  $S \subseteq C$  is clear.  $S$  is inductive, so  $C \subseteq S$ . □

Question: under what conditions do we get "generalized unique readability?" The goal would be to define a function on  $C$  recursively i.e. to have rules for computing  $\bar{h}(x)$  for  $x \in B$  with some rules of computing  $\bar{h}(f(x, y))$  and  $\bar{h}(g(x))$  from  $\bar{h}(x)$  and  $\bar{h}(y)$ .

**Example 1.4. :** Suppose that  $G$  is some additive group, generated from  $B$  (the set of generators),  $h = B \rightarrow H$  where  $(H, \cdot, {}^{-1}, 1)$  a group. When is there an extension  $\bar{h}$  of  $h$  s.th.  $\bar{h} : G \rightarrow H$  is a grouphomomorphism.

- $\bar{h}(0) = 1$
- $\bar{h}(a + b) = \bar{h}(a) \cdot \bar{h}(b)$
- $\bar{h}(-a) = \bar{h}(a)^{-1}$

This is not always possible. **Note:** that it is possible if  $G$  is generated freely by the elements of  $B$  and the set of atoms is independent (one element of  $B$  cannot be generated in finitely many steps by other elements of  $B$ ).

**Definition 1.9. Freely generated set:**  $C$  is freely generated from  $B$  by  $f, g$  if freely generated

- $C$  is generated from  $B$  by  $f, g$
- $f|_{C^2}$  and  $g|_C$  are such that
  1.  $f|_{C^2}$  and  $g|_C$  are one-to-one (injective)
  2.  $\text{ran}(f|_{C^2})$  and  $\text{ran}(g|_C)$  and  $B$  are p.w. disjoint

**Theorem 1.3. Recursion Theorem:**  $C \subseteq U$  freely generated from  $B$  by  $f, g$  and  $V$  a set and  $h : B \rightarrow V$ ,  $F : V^2 \rightarrow V$ ,  $G : V \rightarrow V$  Then  $\exists! \bar{h} : C \rightarrow V$  s.that

- for all  $a$  in  $B$  it holds  $\bar{h}(a) = h(a)$
- for all  $x, y$  in  $C$  it holds
  1.  $\bar{h}(f(x, y)) = F(\bar{h}(x), \bar{h}(y))$
  2.  $\bar{h}(g(x)) = G(\bar{h}(x))$

**Note:** if given conditions are satisfied then  $h$  extends uniquely to a homomorphism

$$(C, f, g) \rightarrow (V, F, G)$$

Before we proof the recursion theorem, we will show how unique readability easily follows from it.

**Note:** Recursion Theorem implies unique readability for propositional formulas. What we need to check is that the Assumptions of recursion theorem are satisfied.

**Claim:** The formula building operations are one-to-one.

*proof of claim.*  $\mathcal{F}_\vee$  is one to one, suppose  $(\alpha \vee \beta) = (\delta \vee \gamma)$  then  $\alpha \vee \beta = \delta \vee \gamma$  And  $\alpha, \delta$  are prop. formulas, so they equal to each other (else one is an initial segment of the other, hence not a prop. fla.) By the same argument we get  $\beta$  is equal to  $\gamma$ . □

**Claim:** Disjointment of ranges

*proof of claim.* • if  $(\alpha \vee \beta) = A$  then  $A$  starts with ( which can not be the case

- if  $(\alpha \vee \beta) = (\gamma \rightarrow \delta)$  then by the same argument  $\alpha$  is  $\gamma$  but  $\vee$  and  $\rightarrow$  are different
  - if  $(\alpha \vee \beta) = (\neg \gamma)$ , then  $\alpha \vee \beta = \neg \gamma$ , so  $\alpha$  would start with a  $\neg$ , -no
- For all other connectives the proof is similar. □

### Proof of the Rec Thm.

$v : C \rightarrow V$  is called acceptable if  $\forall x, y \in C$  acceptable

1. if  $x \in B \cap \text{dom}(v)$  then  $v(x) = h(x)$
2. if  $f(x, y) \in \text{dom}(v)$  then  $x, y \in \text{dom}(v)$  and similarly for  $g$ 
  - $v(f(x, y)) = F(v(x), v(y))$
  - $v(g(x)) = G(v(x))$

And when  $U = \{I_v : v \text{ acceptable}\}$ , we define  $\bar{h} :=$  function w/ graph  $\bigcup I_v$

**Claim 1:**  $\bar{h}$  is a function.

*proof of claim.*

$$S := \{x \in C : \exists \text{at most one } y \text{ with } (x, y) \in \bigcup \Gamma_v\}$$

We want  $S = C$ , we have  $S \subseteq C$ , it is enough to show that  $S$  is inductive.

- $x \in B \cap \text{dom}(v)$  for some  $v$  acceptable.  
then  $v(x) = h(x)$  by 1. also  $x \notin \text{ran}(f|_{C^2})$  and  $x \notin \text{ran}(g|_C)$
- $x, y \in S$  We want  $f(x, y), g(x) \in S$   
there are  $v_1, v_2$  acceptable s.t.  $f(x, y) \in \text{dom}(v_1) \cap \text{dom}(v_2)$

⊠

**Claim 2:**  $\bar{h}$  is acceptable.

*proof of claim.*  $\bar{h} : C \rightarrow V$  by definition. if  $x \in B \cap \text{dom } \bar{h}$  then there is a  $v$  acceptable, s.t.  $x \in \text{dom}(v)$  then  $\bar{h}(x) = v(x) = h(x)$  if  $f(x, y) \in \text{dom } \bar{h}$  then  $f(x, y) \in \text{dom}(v)$  form some  $v$  acceptable. Hence  $x, y \in \text{dom}(v)$  and therefore  $x, y \in \text{dom}(\bar{h})$  and we have

$$\bar{h}(f(x, y)) = v(f(x, y)) = F(v(x), v(y)) = F(\bar{h}(x), \bar{h}(y))$$

⊠

**Claim 3:** The domain of  $\bar{h}$  equals  $C$ .

*proof of claim.* it is enough to show that the domain of  $\bar{h}$  is inductive.  $B \subseteq \text{dom}(\bar{h})$  bc.  $B \subseteq \text{dom}(h)$  where  $h$  is acceptable. Now we need to show closure under  $f, g$ . suppose  $x', y' \in \text{dom}(\bar{h})$  then  $x' \in \text{dom}(v_1)$  for some acceptable  $v_i$  lets assume  $f(x', y') \notin \text{dom}(\bar{h})$  then we extend  $\bar{h}$  to a function with the same graph as  $\bar{h}$ . Then  $\Gamma \cup \{(f(x', y'), F(x', y'))\}$  is the graph of an acceptable function. ⊠

**Claim 4:**  $\bar{h}$  is uniquely constructed

*proof of claim.* Suppose both  $\bar{h}, \bar{\bar{h}}$  work, we show that  $S = \{x \in C : \bar{h}(x) = \bar{\bar{h}}(x)\}$  is the whole set  $C$ . it is enough to show that  $S$  is inductive. Let  $x \in B$  then  $\bar{h}(x) = h(x) = \bar{\bar{h}}(x)$ . Then for  $x, y \in S$

$$\bar{h}(f(x, y)) = F(\bar{h}(x), \bar{h}(y)) = F(\bar{\bar{h}}(x), \bar{\bar{h}}(y)) = \bar{\bar{h}}(f(x, y))$$

$$\bar{h}(g(x)) = G(\bar{h}(x)) = G(\bar{\bar{h}}(x)) = \bar{\bar{h}}(g(x))$$

and  $f(x, y), g(x) \in S$ , therefore  $S$  is inductive. ⊠

## 1.4 SENTENTIAL CONNECTIVES

tautological  
equivalence

$\sim$

$\models$

$\models$

**Definition 1.10. Tautological equivalence relation:** For  $\alpha, \beta$  prop. sent. we define  $\alpha \sim \beta$  iff  $\alpha \models \beta$  (alternative notation:  $\models$ ). This defines an equivalent relation.

**Example 1.5. :**  $A \rightarrow B \models \neg A \vee B$

**Note:** A  $k$ -place boolean function is a function of the form  $f : \{0, 1\}^k \rightarrow \{0, 1\}$  and we define 0, 1 as the 0-place boolean functions.

If  $\alpha$  is a prop. sent. then it determines a  $k$ -place boolean function, where  $k$  is the number of atoms,  $\alpha$  is built up from. If  $\alpha$  is  $(A_1 \vee \neg A_2)$  then  $B_\alpha : \{0, 1\}^2 \rightarrow \{0, 1\}$  and assign its values corresponding a truth value of  $\alpha$ . That is for any TA  $v : \{A_1, A_2\} \rightarrow \{0, 1\}$  we define  $B_\alpha(v(A_1), v(A_2)) = \bar{v}(\alpha)$

**Theorem 1.4. :** If  $\alpha, \beta$  are prop. sent. with at most  $n$  prop. Atoms (combined), then

1.  $\alpha \models \beta$  iff  $\forall x \in \{0, 1\}^n$  it holds  $B_\alpha(x) \leq B_\beta(x)$
2.  $\alpha \models \beta$  iff  $\forall x \in \{0, 1\}^n$  it holds  $B_\alpha(x) = B_\beta(x)$
3.  $\models \alpha$  iff  $\forall x \in \{0, 1\}^n$  it holds  $B_\alpha(x) = 1$

$n$ -ary boolean func.

**Theorem 1.5. Realisation:** Let  $G$  be an  $n$ -ary boolean function for  $n \geq 1$ . Then there is a prop. sent.  $\alpha$  such that.  $B_\alpha = G$ . We say  $\alpha$  realizes  $G$ .



*Proof.* 1. if  $G$  is constantly equal to 0 then set  $\alpha$  to  $A_1 \wedge \neg A_1$ .

2. Otherwise the set of inputs  $\{\vec{x}_1, \vec{x}_2, \dots, \vec{x}_k\}$  for which  $G(\vec{x}_i) = 1$  holds is not empty.  
We denote  $\vec{x}_i = (x_{i1}, x_{i2}, \dots, x_{in})$  and define a matrix  $(x_{ij})_{k \times n}$ . We further set

$$\beta_{ij} = \begin{cases} A_j & \text{iff } x_{ij} = 1 \\ \neg A_j & \text{iff } x_{ij} = 0 \end{cases}$$

**Example:**

$$(x_{ij}) = \begin{pmatrix} 0 & 1 & 0 \\ 1 & 1 & 0 \end{pmatrix} \rightsquigarrow \begin{pmatrix} \neg A_1 & A_2 & \neg A_3 \\ A_1 & A_2 & \neg A_3 \end{pmatrix} = (\beta_{ij})$$

We define  $\gamma_i$  as  $\beta_{i1} \wedge \beta_{i2} \wedge \dots \wedge \beta_{in}$  for  $1 \leq i \leq k$   
and  $\alpha$  as  $\gamma_1 \vee \gamma_2 \vee \dots \vee \gamma_k = \bigvee_{i=1}^k \gamma_i$ . Then  $B_\alpha = G$  is fulfilled.

□

**Note:**  $\alpha$  as constructed in the proof is in the so-called Disjunctive normal form (DNF).

DNF

**Corollary 1.5.** Every prop. sent. is tautologically equivalent to a sentence in DNF

Disjunctive normal form

**Corollary 1.5.**  $\{\neg, \wedge, \vee\}$  is a complete set of logical connectives, i.e. every prop. sent. is tautologically equivalent to a sentence built up from atoms and  $\neg, \wedge, \vee$ .

complete

**Theorem 1.6. :** Both  $\{\neg, \wedge\}$  and  $\{\neg, \vee\}$  are complete.

*Proof.* It's sufficient to show that every  $k$ -place boolean function is realisable by a prop. sent. built up using only  $\neg$  and  $\wedge$ . This is, because  $\alpha \wedge \beta \models \neg(\neg\alpha \vee \neg\beta)$ . We prove this by induction over the number of disjunctions of a prop. sent.  $\alpha$  in DNF. Suppose the statement is true for  $k \leq n$ . For  $n+1$  and  $\alpha = \bigvee_{j=1}^{n+1} \gamma_j$  there exists an  $\alpha' \models \bigvee_{j=1}^n \gamma_j$  and

$$\alpha = \bigvee_{j=1}^{n+1} \gamma_j \models \alpha' \vee \gamma_{n+1} \models \neg(\neg\alpha' \wedge \neg\gamma_{n+1})$$

□

**Note:** We used the observation that, if  $\alpha \models \beta$  and we replace a subsequence of  $\alpha$  by a so called tautological equivalence then the result is also tautologically equivalent to  $\beta$

**Example 1.6.**  $\{\rightarrow, \wedge\}$  is not complete.: Let  $\alpha \in PS$  built up from only  $\rightarrow, \wedge$  from the atoms  $A_1, \dots, A_n$  then we claim

$$A_1 \wedge A_2 \wedge \dots \wedge A_n \models \alpha$$

We can also say  $\{\rightarrow, \wedge\}$  is not complete bc.  $\neg A$  is not tautological equivalent to a sent. built up from  $\rightarrow, \wedge$

*Proof.* Let  $C := \{\alpha \in PS \text{ built up from } \rightarrow, \wedge \text{ and } A_1, \dots, A_n \text{ for which } \bigwedge_{i=1}^n A_i \models \alpha\}$  we want to show that  $C = \{\alpha \in PS \text{ built up from } \rightarrow, \wedge \text{ and } A_1, \dots, A_n\}$

- We have  $\{A_1, A_2, \dots, A_n\} \subseteq C$
- for  $\alpha, \beta \in C$  it holds

$$(1) A_1 \wedge \dots \wedge A_n \models \alpha \rightarrow \beta$$

$$(2) A_1 \wedge \dots \wedge A_n \models \alpha \wedge \beta$$

Therefore  $C$  is closed under the fla. building operations and we have proven our claim. □

**Note:**  $\{\wedge, \vee, \rightarrow, \leftrightarrow\}$  is still not complete.

**Note:** The number of  $n$ -ary boolean functions existing is  $2^{2^n}$ . We define a notation for  $n = 0$ :  $\perp$  (for  $TV = 0$ ) and  $\top$  (for  $TV = 1$ ). We can conclude that  $\{\neg, \rightarrow\}$  and  $\{\rightarrow, \perp\}$  are both complete, it holds  $\neg A \models A \rightarrow \perp$

**Definition 1.11. Satisfiability:** A set of prop. sent.  $\Sigma$  is called **satisfiable** if there exists a TA that satisfies every member of  $\Sigma$ . satisfiable

## 1.5 COMPACTNESS THEOREM

finitely satisfiable

**Theorem 1.7. Compactness Theorem:**  $\Sigma$  is satisfiable iff every finite subset  $\Sigma_0 \subseteq \Sigma$  is satisfiable. (i.e.  $\Sigma$  is finitely satisfiable)

*Proof.* Let  $\Sigma$  be a finitely satisfiable set of prop. sent. Outline of the proof:

1. extend  $\Sigma$  to a maximal finitely satisfiable set  $\Delta$  of prop. sent.

2. construct a truth assignment using  $\Delta$

1. Let  $\alpha_1, \alpha_2, \dots$  be an enumeration of all prop. sent. and define  $\Delta_n$  inductively by  $\Delta_0 := \Sigma$

$$\Delta_{n+1} := \begin{cases} \Delta_n \cup \{\alpha_{n+1}\} & \text{if satisfiable} \\ \Delta_n \cup \{\neg\alpha_{n+1}\} & \text{otherwise} \end{cases}$$

**Claim:**  $\Delta_n$  is finitely satisfiable for each  $n$

*proof of claim.* By regular induction over  $n$ .  $\Delta_0$  is finitely satisfiable. Let us assume  $\Delta_n$  is finitely satisfiable. If  $\Delta_{n+1} = \Delta_n \cup \{\alpha_{n+1}\}$  then we are finished. Otherwise let  $\Delta' \subseteq \Delta_n$  be a finite set that  $\Delta' \cup \{\alpha_{n+1}\}$  is not satisfiable. It holds  $\Delta' \models \neg\alpha_{n+1}$ . We assume that  $\Delta_n \cup \{\neg\alpha_{n+1}\}$  is not finitely satisfiable. Then there exists a finite subset  $\Delta'' \subseteq \Delta_n$  such that  $\Delta'' \cup \{\neg\alpha_{n+1}\}$  is (finite and) not satisfiable. It therefore holds  $\Delta'' \models \alpha_{n+1}$ . But  $\Delta' \cup \Delta''$  is a finite subset of  $\Delta_n$  and by above observations  $\Delta' \cup \Delta'' \models \alpha_{n+1}$  and  $\Delta' \cup \Delta'' \models \neg\alpha_{n+1}$ . A contradiction to the assumption that  $\Delta_n$  is finitely satisfiable.  $\square$

We set  $\Delta := \bigcup_{i \in \mathbb{N}} \Delta_i$  and get

(a)  $\Sigma \subseteq \Delta$

(b) (Maximality): for every prop. sent.  $\alpha$  it holds  $\alpha \in \Delta$  or  $\neg\alpha \in \Delta$

(c) (Satisfiability):  $\Delta$  is finitely satisfiable. For every finite subset there exists a  $\Delta_n$  which is a superset.

2. Let  $\nu$  be a TA for the prop. atoms  $A_1, A_2, \dots$  such that  $\nu(A) = 1$  iff  $A \in \Delta$

**Claim:** For every prop. sent.  $\varphi$  it holds  $\bar{\nu}(\varphi) = 1$  iff  $\varphi \in \Delta$ .

*proof of claim.* Let  $S = \{\varphi \in PS \text{ s.t. } \bar{\nu}(\varphi) = 1 \text{ iff } \varphi \in \Delta\}$ .

•  $PS \supseteq S$  is clear.

•  $PS \subseteq S$

(a)  $\{A_1, A_2, \dots\} \subseteq S$  by definition of  $\nu$

(b) closure under  $\neg$ : Let  $\varphi \in S$  then we get by maximality and satisfiability of  $\Delta$ :

$$\begin{aligned} \bar{\nu}(\neg\varphi) &= 1 \\ \text{iff } \bar{\nu}(\varphi) &= 0 \\ \text{iff } \varphi &\notin \Delta \\ \text{iff } (\neg\varphi) &\in \Delta \end{aligned}$$

closure under  $\rightarrow$ : Let  $\varphi_1, \varphi_2 \in S$  similarly

$$\begin{aligned} \bar{\nu}(\varphi_1 \rightarrow \varphi_2) &= 0 \\ \text{iff } \bar{\nu}(\varphi_1) &= 1 \text{ and } \bar{\nu}(\varphi_2) = 0 \\ \text{iff } \varphi_1 &\in \Delta \text{ and } \varphi_2 \notin \Delta \\ \text{iff } (\varphi_1 \rightarrow \varphi_2) &\notin \Delta \end{aligned}$$

The closure under the other fla. building operations are similar.  $\square$

By this claim  $\bar{\nu}$  satisfies  $\Sigma$ .  $\square$

**Corollary 1.7.** If  $\Sigma \models \tau$  then there exists a finite subset  $\Sigma' \subseteq \Sigma$  s.t.  $\Sigma' \models \tau$

*Proof.* Recall:  $\Sigma \models \tau$  iff  $\Sigma \cup \{\neg\tau\}$  is not satisfiable. Suppose  $\Sigma \models \tau$  but no finite subset does.

Then  $\forall \Sigma' \subseteq \Sigma$  finite  $\Sigma' \cup \{\neg\tau\}$  is satisfiable. By the compactness theorem  $\Sigma \cup \{\neg\tau\}$  is satisfiable which is a contradiction to  $\Sigma \models \tau$ .  $\square$

**Note:** Theorem 1.7 and Corollary 1.7 are equivalent.

## CHAPTER 2

# Predicate - / first order logic

**Definition 2.1. A First order Language:** consists of infinitely many distinct symbols such that no symbol is a proper initial segment of another symbol and the symbols are divided into 2 groups:

1. logical symbols logical symbols  
(These elements have a fixed meaning and the equivalence symbol  $=$  is optional)  
 $(, ), \neg, \rightarrow, v_1, v_2, \dots, =$
2. parameters parameters
  - quantifier symbol:  $\forall$  (the range is subject of interpretation)
  - predicate symbols: for every  $n > 0$  we have a set of  $n$ -ary predicates  $P$
  - constant symbols: Some set of constants (could also be  $\emptyset$ )
  - function symbols: for every  $n > 0$  we have a set of  $n$ -ary function symbols

**Note:**

- We could drop constants and instead introduce 0-ary function symbols
- to specify language we need to specify the parameters and say if  $=$  is included
- In the book [1] they assume that some  $n$ -place predicate symbol is present for some  $n$ .

**Example 2.1. :**

- $\mathcal{L}_{\text{set}} = \{\in\}$ ,  $=$  is included and the binary predicate symbol  $\in$  "element in"
- $\mathcal{L}_{\text{arith}} = \{<, 0, S, E, +, \cdot\}$ 
  - $=$  included
  - $<$  is a binary rel. symbol
  - $0$  is a constant
  - $S$  is a unary function symbol
  - $E$  exponentiation function symbol
  - $+, \cdot$  binary function symbols
- $\mathcal{L}_{\text{ring}} = \{=, +, \cdot, -, 0, 1\}$ 
  - $=$  included
  - $0, 1$  are constants
  - $-$  is a unary function symbol (additive inverse)
  - $+, \cdot$  binary function symbols

## 2.1 FORMULAS

**Definition 2.2. Expression:** An **expression** is any finite sequence of symbols. There exist two kinds of expressions that makes sense "grammatically"

- Terms:
- points to an object
  - they are built up from variables and constants using function symbols

- Formulas:
- They express assertions about objects,
  - they are built up from atomic formulas
  - atomic formulas these are built up from terms using predicate symbols and  $=$ , if included

**Definition 2.3. Term Building Operations:** For every  $n > 0$  and for every  $n$ -place function symbol  $f$  let  $\mathcal{F}_f$  be an  $n$ -place term building operation, that is  $\mathcal{F}_f(t_1, \dots, t_n) := ft_1, \dots, t_n$  (polish notation for  $f(t_1, \dots, t_n)$ ). The Set of terms we then define as the set of expressions that are built up from variables and constants by applying the term building operations finitely many times.

**Example 2.2. :** Let  $\mathcal{L} = \mathcal{L}_{arith}$  then the set of terms will contain  $0, v_{42}, S0, SSS0, Sv_1, +SOv_1$

**Definition 2.4. Atomic formula:** Any expression of the form

$$= t_1 t_2 \text{ or } P t_1, \dots, t_n, \text{ where } t_1, \dots, t_n \text{ are terms and } P \text{ is an } n\text{-ary predicate symbol}$$

**Note:** Atomic formulas are not defined inductively.

**Example 2.3. :**  $cont. = v_1 v_{42}, < S0 S S0$  are atomic formulas, but  $\neg = v_1 v_{42}$  is not.

**Definition 2.5. Formulas:** We define  $\varepsilon_{\neg}, \varepsilon_{\rightarrow}, Q_i$  to be the fla. building operations, defined as follows  $\varepsilon_{\neg}(\alpha) := (\neg\alpha)$ ,  $\varepsilon_{\rightarrow} := (\alpha \rightarrow \beta)$  and  $Q_i(\gamma) := \forall v_i \gamma$ . The set of formulas is the set of expressions built up from atomic formulas by applying the fla. building operations finitely many times.

**Example 2.4. :**  $cont. \forall v_1 (= Sv_1 0)$  is a formula we get by applying  $Q_1$  on the atomic formula  $= Sv_1 0$ .

### Free variables

$\exists$  quantifier **Example 2.5. :** We introduce the  $\exists$  quantifier by defining  $\exists y \alpha$  means  $\neg \forall y \neg \alpha$ .

bounded variable "Every non-zero natural number is a succesor"  $\forall x (x \neq 0 \rightarrow \exists y S(y) = x)$  is different then "if a number is not 0, then it is a succesor"  $x \neq 0 \rightarrow \exists y S(y) = x$ .  $x$  occurs bounded in the first formula, for the latter  $x$  occurs free in the fla.

If you have an expression without free variables, it is either true or false, on the other hand if a variable occurs free in a formula, the truth value of it depends on the variable itself.

**Definition 2.6. Free variables:** Let  $x$  be a variable.  $x$  occurs **free** in  $\varphi$  is defined inductively as follows:

1. If  $\varphi$  is an atomic fla. then  $x$  occurs **free** in  $\varphi$  iff  $x$  occurs in  $\varphi$
2. If  $\varphi = (\neg\alpha)$  then  $x$  occurs free in  $\varphi$  iff  $x$  occurs free in  $\alpha$
3. If  $\varphi = (\alpha \rightarrow \beta)$  then  $x$  occurs free in  $\varphi$  iff  $x$  occurs free in  $\alpha$  or  $\beta$
4. If  $\varphi = \forall v_i \alpha$  then  $x$  occurs free in  $\varphi$  iff  $x$  occurs free in  $\alpha$  and  $x \neq v_i$

sentence A formula  $\alpha$  is called a sentence, if no variable occurs free in  $\alpha$

**Note:** The above definition makes sense thanks to the recursion theorem. define the function  $h$  on the set of atoms:  $h(\alpha) =$  the set of var occ in fla  $\alpha$ , which is the set of all variables  $v_i$  that occur free in  $\alpha$ . we now want to extend  $h$  to  $\bar{h}$ , which is the set of all formulas.

- $\bar{h}(\neg\alpha) = \bar{h}(\alpha)$
- $\bar{h}(\alpha \rightarrow \beta) = \bar{h}(\alpha) \cup \bar{h}(\beta)$
- $\bar{h}(Q_i(\alpha)) = \bar{h}(\alpha) \setminus \{v_i\}$

We say  $x$  occurs free in  $\alpha$  iff  $x \in \bar{h}(\alpha)$ .

**Note:** We will now use  $\neg, \wedge, \vee, \rightarrow, \leftrightarrow, \exists v_i$  (all can be expressed in terms of  $\neg, \rightarrow, Q_i$ .) We will sometimes drop the  $(, )$  and not always be using polish notation.

## 2.2 SEMANTICS OF FIRST ORDER LOGIC

The equivalent scheme to our TA in predicate logic. The meaning of formulas is given by *structures*, which also determine the scope of the quantifier  $\forall$ , the meaning of all parameters.

**Definition 2.7. structure:** A structure  $\mathcal{A}$  for a first order language  $\mathcal{L}$  is a non-empty set  $A$  called **universe** or **underlying set** of  $\mathcal{A}$  together with an interpretation of each parameters of  $\mathcal{L}$  i.e.

- $\forall$  ranges over the universe  $A$
- for an  $n$ -ary pred. symbol  $P \in \mathcal{L}$  its interpretation  $P^{\mathcal{A}}$  is a subset of  $A^n$  interpretation
- for a constant  $c \in \mathcal{L}$  its interpretation  $c^{\mathcal{A}}$  is an element of  $A$
- for an  $n$ -ary function symbol  $f \in \mathcal{L}$  its interpretation  $f^{\mathcal{A}}$  is a total function

$$f^{\mathcal{A}} : A^n \rightarrow A$$

**Note:**  $A \neq \emptyset$ , and all functions  $f^{\mathcal{A}}$  are total.

**Example 2.6. :** Let  $\mathcal{L} = \{\in\}$  where  $\in$  is a binary relation " An example of an  $\mathcal{L}$  structure is  $(\mathbb{N}, \in^{\mathbb{N}})$  where  $\in^{\mathbb{N}} = \{(x, y) \in \mathbb{N}^2 : x < y\}$

**Definition 2.8. Assignment:** Let  $\varphi$  be a  $\mathcal{L}$ -fla. and  $\mathcal{A}$  a  $\mathcal{L}$ -structure. Let  $V$  be the set of all variables in  $\mathcal{L}$  and  $s : V \rightarrow A$  an assignment. We define the extention  $\bar{s}$  of  $s$  to the set of all  $\mathcal{L}$ -terms by assignment

- if  $x \in V$  then  $\bar{s}(x) := s(x)$
- for  $c \in \mathcal{L}$  a constant symbol, then  $\bar{s}(c) := c^{\mathcal{A}}$
- for  $t_1, \dots, t_n$   $\mathcal{L}$ -terms and  $f \in \mathcal{L}$  an  $n$ -ary function symbol, then

$$\bar{s}(f t_1 \dots t_n) := f^{\mathcal{A}}(\bar{s}(t_1), \dots, \bar{s}(t_n))$$

**Note:** in the previous definition point 3. for  $n = 1$  yields a commutative diagram.

**Theorem 2.1. :** For any given assignment  $s$  there exists a unique extention  $\bar{s}$  as in the previous definition.

*Proof.* will follow from recursion theorem and unique decomposition of terms. □

## Definition of truth

$\models_{\mathcal{A}}$  **Definition 2.9. Satisfy:** We define ' $\mathcal{A}$  satisfies  $\varphi$  with  $s$ ' and write  $\mathcal{A} \models \varphi[s]$  or  $\models_{\mathcal{A}} \varphi[s]$  inductively over the complexity of the formula  $\varphi$

1. if  $\varphi$  is atomic:
  - $\mathcal{A} \models t_1, t_2 [s]$  iff  $\bar{s}(t_1) = \bar{s}(t_2)$
  - $\mathcal{A} \models Pt_1, \dots, t_n [s]$  iff  $(\bar{s}(t_1), \dots, \bar{s}(t_n)) \in P^{\mathcal{A}}$
2. suppose  $\mathcal{A} \models \varphi[s]$  and  $\mathcal{A} \models \psi[s]$  are defined, then
  - $\mathcal{A} \models \neg\varphi[s]$  iff  $\mathcal{A} \not\models \varphi[s]$
  - $\mathcal{A} \models \varphi \rightarrow \psi[s]$  iff  $\mathcal{A} \models \psi[s]$  or  $\mathcal{A} \not\models \varphi[s]$
  - $\mathcal{A} \models \forall x\varphi[s]$  iff for all  $a \in A$   $\mathcal{A} \models \varphi[s(x|a)]$  where

$$s(x|a)(v) = \begin{cases} s(v) & \text{if } v \neq x \\ a & \text{if } v = x \end{cases}$$

**Example 2.7. :**  $\mathcal{L} = \{\forall, \leq, S, 0\}$  a  $\mathcal{L}$ -structure then could be  $\mathcal{N} = (\mathbb{N}, \leq^{\mathcal{N}}, S^{\mathcal{N}}, 0^{\mathcal{N}})$  together with an assignment  $s : v_n \mapsto n - 1$  then:

- $s(v_1) = 0$
- $\bar{s}(0) = 0^{\mathcal{N}}$  (a constant is always mapped to its realisation, the interpretation of constant 0 in the structure  $\mathcal{N}$ )
- $\bar{s}(Sv_1) = S^{\mathcal{N}}(\bar{s}(v_1)) = S^{\mathcal{N}}(0) = 1$
- $\mathcal{N} \models \forall v_1 (S(v_1) \neq v_1) [s]$   
 iff for all  $a \in \mathbb{N}$  we have that  $\mathcal{N} \models (S(v_1) \neq v_1)[s(v_1|a)]$   
 iff ...  
 iff for all  $a \in \mathbb{N}$  we have  $S^{\mathcal{A}}(a) \neq a$ , which is true in our structure of the natural numbers.
- Is it true in  $\mathcal{N}$  that  $\mathcal{N} \models S(0) \leq S(v_1) [s]$ ? Yes because

$$\begin{aligned} \mathcal{N} \models S(0) \leq S(v_1) [s] \\ \text{iff } 1 \leq 1 \end{aligned}$$

**Note:** To know wheter  $\mathcal{A} \models \varphi[s]$  it suffices to know where  $s$  maps the variables that are free in  $\varphi$

**Theorem 2.2. :** Suppose  $s_1, s_2 : V \rightarrow A$  agree on all variables that occur free in  $\varphi$  then

$$\mathcal{A} \models \varphi[s_1] \text{ iff } \mathcal{A} \models \varphi[s_2]$$

*Proof.* By complexity of  $\varphi$

1. if  $\varphi$  is  $Pt_1, \dots, t_n$  note: any var that occur in  $\varphi$  occur free in  $\varphi$ , so  $s_1, s_2$  agree on all variables that occur in the terms  $t_1, \dots, t_n$ .

So we Claim: for  $t$  a term,  $s_1, s_2$  assignments that agree on all variables of  $t$  then  $\bar{s}_1(t) = \bar{s}_2(t)$

*proof of claim.* By complexity of  $t$

- $t = v_m$  then  $\bar{s}_1(t) = s_1(v_m) = s_2(v_m) = \bar{s}_2(t)$
- $t = c$  then  $\bar{s}_1(t) = c^{\mathcal{A}} = \bar{s}_2(t)$
- $t = ft_1 \dots t_n$  inductively, assume  $\bar{s}_1(t_i) = \bar{s}_2(t_i)$  for all  $1 \leq i \leq n$  then TODO

□

2. if  $\varphi$  is  $t_1, t_2$  is similar

3. if  $\varphi$  is  $\neg\alpha$  then  $\mathcal{A} \models \neg\alpha[s_1]$  iff  $\mathcal{A} \not\models \alpha[s_1]$  iff  $\mathcal{A} \models \alpha[s_2]$  iff  $\mathcal{A} \models \neg\alpha[s_1]$

4. if  $\varphi$  is  $\alpha \rightarrow \beta$  then  $\mathcal{A} \models \alpha \rightarrow \beta [s_1]$  iff .. or .. iff for  $s_2$  iff ... or ..
5. if  $\varphi$  is  $\forall x \alpha$  then the assumption is that  $s_1, s_2$  .. the free variables of  $\alpha$  are the free variables of  $\varphi$  except for  $x$ . but because  $s_1(x|a) = s_2(x|a)$  they both agree on all free variables of  $\alpha$ .

$$\begin{aligned} \mathcal{A} \models \forall x \varphi [s_1] &\text{ iff for all } a \in A \mathcal{A} \models \varphi [s_1(x|a)] \\ &\text{ iff for all } a \in A \mathcal{A} \models \varphi [s_2(x|a)] \\ &\text{ iff } \mathcal{A} \models \forall x \varphi [s_2] \end{aligned}$$

□

Notation:  $\mathcal{A} \models \varphi$  means that all free variables of  $\varphi$  are among  $v_1, \dots, v_n$  and  $\mathcal{A} \models \varphi [s]$  whenever  $s(v_i) = a_i$  for all  $1 \leq i \leq n$ .

**Corollary 2.2.** If  $\sigma$  is a sentence then  $\mathcal{A} \models \varphi [s]$  for all  $s : V \rightarrow A$  or  $\mathcal{A} \models \varphi [s]$  for all  $s : V \rightarrow A$ .

Notation:  $\mathcal{A} \models \sigma$  "σ is true in  $\mathcal{A}$ ,  $\mathcal{A}$  is a model of  $\sigma$  or  $\sigma$  holds in  $\mathcal{A}$ .

**Note:** If  $\sigma$  is a sentence then we can not have  $\mathcal{A} \models \sigma$  and  $\mathcal{A} \models \sigma$  because  $A \neq \emptyset$ .

**Definition 2.10. Model:**  $\mathcal{A}$  is a model of a set of sentences  $\Sigma$  iff for every sentence  $\sigma \in \Sigma$  it holds  $\mathcal{A} \models \sigma$

**Example 2.8. :**  $\mathcal{L} = \{0, 1, +, -, \cdot\}$  A realisation could be  $\mathcal{R} = (\mathbb{R}, 0, 1, +, -, \cdot)$  or  $\mathcal{C} = (\mathbb{C}, 0, 1, +, -, \cdot)$  then the sentence  $\sigma : \exists x(x \cdot x = -1)$  then  $\mathcal{R} \models \sigma$  but  $\mathcal{C} \models \sigma$

**Note:**  $\wedge, \vee, \leftrightarrow, \exists$  work as expected. That is  $\mathcal{A} \models (\alpha \wedge \beta) [s]$  iff  $\mathcal{A} \models \alpha [s]$  and  $\mathcal{A} \models \beta [s]$   
 $\mathcal{A} \models (\alpha \vee \beta) [s]$  iff  $\mathcal{A} \models \alpha [s]$  or  $\mathcal{A} \models \beta [s]$   $\mathcal{A} \models \exists x \alpha [s]$  iff  $\mathcal{A} \models \neg \forall x \neg \alpha [s]$   
 iff  $\mathcal{A} \models \forall x \neg \alpha [s]$   
 iff it is not true that for all  $a \in A$   $\mathcal{A} \models \neg \alpha [s(x|a)]$   
 iff there is  $a \in A$  such that  $\mathcal{A} \models \alpha [s(x|a)]$

## 2.3 LOGICAL IMPLICATION

Let  $\Gamma$  be a set of  $\mathcal{L}$ -formulas,  $\varphi$  a  $\mathcal{L}$ -formula.

**Definition 2.11. Logical implication:**  $\Gamma \models \varphi$  "Γ logically implies  $\varphi$ " if for every  $\mathcal{L}$ -structure  $\mathcal{A}$  and for every  $s : V \rightarrow A$   
 if  $\mathcal{A} \models \gamma [s]$  for every  $\gamma \in \Gamma$  then  $\mathcal{A} \models \varphi [s]$

**Definition 2.12. Logical equivalence:**  $\varphi, \psi$  are called logically equivalent if  $\varphi \models \psi$  and  $\psi \models \varphi$ .

**Definition 2.13. Valid:**  $\varphi$  is called valid iff  $\models \varphi$  i.e.  $\emptyset \models \varphi$  i.e. for every  $\mathcal{L}$ -structure  $\mathcal{A}$  and every  $s : V \rightarrow A$  it is  $\mathcal{A} \models \varphi [s]$

**Example 2.9. :**

1.  $\forall x_1 P x_1 \models P x_2$   
 Suppose  $\mathcal{A} \models \forall x_1 P x_1 [s]$ . then for all  $a \in A$  it is  $\mathcal{A} \models P x_1 [s(x_1|a)]$  in particular,  $a \in P^{\mathcal{A}}$  for  $a = s(x_2)$
2.  $\forall P x_2 \models \forall x_1 P x_1$   
 We need a counterexample to  $\forall P x_2 \models \forall x_1 P x_1$ . Let  $A = \{a_1, a_2\}$   $s(x_2) = a_1$  and  $P^{\mathcal{A}} = \{a_1\}$  then  $\mathcal{A} \models P x_2 [s]$ .
3. Is the following valid?  $\models \exists x(Px \rightarrow \forall yPy)$  yes
4.  $\Gamma, \alpha \models \varphi$  iff  $\Gamma \models \alpha \rightarrow \varphi$ . (on next problem set, quite important)

## 2.4 DEFINABILITY IN A STRUCTURE

**Definition 2.14. definability in a structure:** We say that a general  $n$ -ary relation  $P$  on  $A$  (we will just call it  $P$ , it does not have to be in the language) is definable in  $\mathcal{A}$ , if there is a  $\mathcal{L}$ -formula  $\varphi$  with free variables among  $\{v_1, \dots, v_n\}$  such that

$$P = \{(a_1, \dots, a_n) : \mathcal{A} \models \varphi[a_1, \dots, a_n]\}$$

We also say that  $\varphi$  defines  $P$  in the structure  $\mathcal{A}$ .

**Example 2.10. :**

1.  $x = x$  would define the entire universe.
2.  $\neg x = x$  would define the empty set.

**Example 2.11. :**

1. TODO
2.  $\mathcal{R} = (\mathbb{R}, 0, 1, +, -, \cdot)$   $Q$ : is  $[0, \infty)$  definable in  $\mathcal{R}$  Yes because  $\exists y(y \cdot y = x)$  Indeed we can even define the  $\leq$  relation on  $\mathbb{R}^2$  by  $x \leq z \Leftrightarrow \exists y(x + y \cdot y = z)$

**Definition 2.15. definability of classes of structures:** Let  $\Sigma$  be a set of sentences.  $\tau$  a sentence. We will say that the class of models of  $\Sigma$  is the class  $Mod\Sigma = \{\mathcal{A} : \mathcal{A} \models \Sigma\}$ . Let  $K$  be a class of structures. We are going to call  $K$  an elementary class (EC) if there is a single sentence  $\tau$  such that  $Mod\tau = K$   $K$  is an elementary class in the wider sense (EC $\Delta$ ) if there is a set of sentences  $\Sigma$  such that  $Mod\Sigma = K$

**Example 2.12. :**  $\mathcal{L} = \{0, 1, +, \cdot\}$   $\tau$  is a sentence that expresses the field axioms (the unary inverse functions are not in our language but are definable.)  $Mod\tau$  is the class of all the fields, which is EC. the class of all fields of characteristic 0. Let  $\sigma_p : \neg(1 + \dots + 1 = 0)$  then  $\Sigma = \{\tau\} \cup \{\sigma_p : p \in \mathbb{P}\}$  yields  $Mod\Sigma$  is the class of fields with characteristic 0, therefore EC $\Delta$ , we will later see that it is not EC.

**Example 2.13. :** Let  $E$  be a binary relation,  $\mathcal{L} = \{E\}$  then a graph is a realisation  $\mathcal{G} = (V, E^{\mathcal{G}})$  such that  $V \neq \emptyset$ ,  $E^{\mathcal{G}}$  is irreflexive and symmetric. By definition the universe is not empty, we still have to check irreflexive and symmetric.

- irreflexive:  $\forall x(\neg xEx)$
- symmetric:  $\forall x\forall y(xEy \rightarrow yEx)$

We take  $\tau$  to be  $\forall x\forall y((\neg xEx) \wedge (xEy \rightarrow yEx))$  Then  $Mod\tau$  is the class of all graphs and is EC Note: the class of all finite graphs is neither EC nor EC $\Delta$ . proof later.

We want to have some notion that tells us when two graphs are the same or at least similar.

## 2.5 HOMOMORPHISMS OF STRUCTURES

**Definition 2.16. Homomorphism:** Suppose that  $\mathcal{A}, \mathcal{B}$  are two  $\mathcal{L}$ -structures. then a Homomorphism of  $\mathcal{A}$  into  $\mathcal{B}$  is a map  $h : A \rightarrow B$  that satisfy the below conditions

- for every  $n$ -ary predicate  $P \in \mathcal{L}$  it is  $(a_1, \dots, a_n) \in P^{\mathcal{A}}$  iff  $(h(a_1), \dots, h(a_n)) \in P^{\mathcal{B}}$  (this def. a strong Homomorphism, other textbooks maybe only require  $\rightarrow$  direction)
- for every  $n$ -ary function  $f \in \mathcal{L}$  and for all  $\underline{a} = (a_1, \dots, a_n) \in A^n$  it holds  $h(f^{\mathcal{A}}(\underline{a})) = f^{\mathcal{B}}(h(a_1), \dots, h(a_n))$
- for every constant symbol  $c \in \mathcal{L}$  it is  $h(c^{\mathcal{A}}) = c^{\mathcal{B}}$  (could also skip this if we consider constants as 0-ary functions)



**Note:** Intuitively a Homomorphism of  $\mathcal{A}$  into  $\mathcal{B}$  is a map  $A \rightarrow B$  that preserve all function and relation symbols in some sense, (imp: not the definable relations)

**Definition 2.17. Isomorphism:**

- $h : A \rightarrow B$  is called isomorphism of  $\mathcal{A}$  into  $\mathcal{B}$  if  $h$  is a Homomorphism and injective (in other textbooks: an isomorphic embedding of  $\mathcal{A}$  into  $\mathcal{B}$ )
- $h : A \rightarrow B$  is called isomorphism of  $\mathcal{A}$  onto  $\mathcal{B}$  if  $h$  is a Homomorphism and bijective  $A \rightarrow B$
- $\mathcal{A}$  and  $\mathcal{B}$  are called isomorphic if there is an isomorphism of  $\mathcal{A}$  onto  $\mathcal{B}$

isomorphic

**Note:**

**Example 2.14. :**  $\mathcal{L} = \{+, \cdot\}$   $\mathcal{N} = (\mathbb{N}, +^{\mathcal{N}}, \cdot^{\mathcal{N}})$  and  $\mathcal{B} = (B, +^{\mathcal{B}}, \cdot^{\mathcal{B}})$  where  $B = \{0, 1\}$

and  $\begin{array}{c|cc} +^{\mathcal{B}} & e & 0 \\ \hline e & e & 0 \\ 0 & 0 & e \end{array} \quad \begin{array}{c|cc} \cdot^{\mathcal{B}} & e & 0 \\ \hline e & e & e \\ 0 & e & 0 \end{array}$  let  $h : \mathbb{N} \rightarrow B$  a Homomorphism?  $h(n) = \begin{cases} e & \text{if } n \text{ is even} \\ 0 & \text{else} \end{cases}$

need at first that  $h(m+n) = h(m) +^{\mathcal{B}} h(n)$  and  $h(m \cdot n) = h(m) \cdot^{\mathcal{B}} h(n)$ . it is indeed a Homomorphism.

**Definition 2.18. Substructure:** Suppose we have two  $\mathcal{L}$  structures and  $A \subseteq B$  then  $\mathcal{A}$  is a substructure of  $\mathcal{B}$  (notation:  $\mathcal{A} \subseteq \mathcal{B}$  or we might say  $\mathcal{B}$  is an extension of  $\mathcal{A}$ ) if

- for every  $n$ -ary relation  $P^{\mathcal{A}} = P^{\mathcal{B}}|_A$
- for every  $n$ -ary function  $f^{\mathcal{A}} = f^{\mathcal{B}}|_A$
- for every constant symbol  $c$  in  $\mathcal{L}$  it is  $c^{\mathcal{A}} = c^{\mathcal{B}}$

**Example 2.15. :**  $\mathcal{L} = \{\leq\}$  then  $\mathcal{N} = (\mathbb{N}, \leq)$  and  $\mathcal{P} = (\mathbb{N}^+, \leq^{\mathcal{P}})$  where  $\leq^{\mathcal{P}}$  is the restriction of  $\leq$  to the positive natural numbers.  $\mathcal{P} \subseteq \mathcal{N}$  and there exists a isomorphic embedding  $id : \mathbb{N}^+ \rightarrow \mathbb{N}$  from  $\mathcal{P}$  into  $\mathcal{N}$  They are even isomorphic ( $h : \mathbb{N} \rightarrow \mathbb{N}^+, h(n) = n+1$ ) so in fact  $\mathcal{P} \cong \mathcal{N}$ .

**Example 2.16. :**  $(\mathbb{Q}, +) \subseteq (\mathbb{C}, +)$

**Note:** If  $\mathcal{A} \subseteq \mathcal{B}$  then in particular  $\mathcal{A}$  is closed under all constant and functions in  $\mathcal{B}$  So suppose that  $\mathcal{B}$  is a substructure and  $A \subseteq B$  and  $A \neq \emptyset$  and  $A$  is closed under  $f^{\mathcal{B}}, c^{\mathcal{B}}$  Can then  $A$  be made into a substructure  $\mathcal{A}$  of  $\mathcal{B}$ .  $f^{\mathcal{A}}$  would be the restriction of  $f^{\mathcal{B}}$  to  $A^n$ , constants  $c^{\mathcal{A}} = c^{\mathcal{B}}$  and if  $P \in \mathcal{L}$  is an  $n$ -ary predicate then  $P^{\mathcal{A}}$  should be  $P^{\mathcal{B}} \cap A^n$ . If  $\mathcal{L}$  has no const. or function symbols then any subset can be made into a substructure of a structure on  $\mathcal{L}$ .

Our next question will be: what is the relation of the above notions with truth and satisfiability The answer will be given by the so called Homomorphism theorem.

**Theorem 2.3. Homomorphism theorem:**  $h$  homomorphism of  $\mathcal{A}$  into  $\mathcal{B}$ ,  $s : V \rightarrow A$  then

1. for all terms  $t$  it is  $h(\overline{s(t)}) = \overline{(h \circ s)(t)}$
2.  $\varphi$  a fla. that is quantifier free and does not include  $=$  then  $\mathcal{A} \models \varphi[s]$  iff  $\mathcal{B} \models \varphi[h \circ s]$
3. if  $h$  is additionally injective then we can drop the requirement "no  $=$ ".
4. if  $h$  is homomorphism of  $\mathcal{A}$  onto  $\mathcal{B}$  then we can drop the requirement "q.f." in (b)

*Proof.* 1. problem set

2.
  - $\varphi : Pt$  then  $\mathcal{A} \models Pt[s]$  iff  $\overline{s(t)} \in P^{\mathcal{A}}$  iff  $h(\overline{s(t)}) \in P^{\mathcal{B}}$  iff  $\overline{(h \circ s)(t)} \in P^{\mathcal{B}}$  iff  $\mathcal{B} \models Pt[h \circ s]$
  - $\varphi : \neg\psi$   $\mathcal{A} \models \neg\psi[s]$  iff  $\mathcal{A} \not\models \psi[s]$  iff  $\mathcal{A} \models \psi[s]$  iff
  - $\varphi : \psi \rightarrow \alpha$
3.  $\mathcal{A} \models t_1 t_2[s]$  iff  $\overline{s(t_1)} = \overline{s(t_2)}$  iff  $h(\overline{s(t_1)}) = h(\overline{s(t_2)})$  iff (by (a))  $\overline{(h \circ s)(t_1)} = \overline{(h \circ s)(t_2)}$  iff  $\mathcal{B} \models t_1 t_2[h \circ s]$

4.  $\varphi \forall s : V \rightarrow A \mathcal{A} \models \varphi[s]$  iff  $\mathcal{B} \models \varphi[h \circ s]$ , want  $\mathcal{A} \models \forall x \varphi[s]$  iff  $\mathcal{B} \models \forall x \varphi[h \circ s]$  1.  $\mathcal{B} \models \forall x \varphi[(h \circ s)]$  iff for all  $s : V \rightarrow A$ ,  $a \in A$  (req. surjectivity) it is  $\mathcal{B} \models \varphi[(h \circ s)(x|h(a))]$  iff  $\mathcal{B} \models \varphi[h \circ (s(x|a))]$  iff (inductive assumption)  $\mathcal{A} \models \varphi[s(x|a)]$  because  $a$  was arbitrary it is  $\mathcal{A} \models \forall x \varphi[s]$  2. Suppose  $\mathcal{B} \models \forall x \varphi[(h \circ s)]$  then there exists a  $b \in B$  such that  $\mathcal{B} \models \neg \varphi[(h \circ s)(x|b)]$  by surjectivity we can find  $a \in A$  such that  $h(a) = b$  and it is  $\mathcal{B} \models \neg \varphi[(h \circ s)(x|h(a))]$  By the inductive assumption  $\mathcal{A} \models \neg \varphi[s(x|a)]$  and  $\mathcal{A} \models \forall x \varphi[s]$   $\square$

**Note:**  $\mathcal{A} \cong \mathcal{B}$  then  $\mathcal{A}$  and  $\mathcal{B}$  satisfy exactly the same sentences.

**Definition 2.19. elementarily equivalent:**  $\mathcal{A}$  and  $\mathcal{B}$  are called elementarily equivalent ( $\mathcal{A} \equiv \mathcal{B}$ ) if  $\mathcal{A}$  and  $\mathcal{B}$  satisfy the same sentences.

**Note:** If  $\mathcal{A} \cong \mathcal{B}$  implies  $\mathcal{A} \equiv \mathcal{B}$  The converse is not true. For instance DLO (dense linear order) w/o endpoints is complete, so two structures on DLO are equivalent  $(\mathbb{Q}, <) \equiv (\mathbb{R}, <)$  but they are not isomorphic because the universes have different cardinality.

**Example 2.17. :**  $\mathcal{N} = (\mathbb{N}, \leq)$  and  $\mathcal{P} = (\mathbb{N}^{>0}, \leq)$   $h : n \mapsto n - 1 : \mathcal{P} \rightarrow \mathcal{N}$  isom. so in part  $\mathcal{N} \equiv \mathcal{P}$ . but  $id : \mathcal{P} \rightarrow \mathcal{N}$  is only isom embedding, so for example  $\forall y (x \neq y \rightarrow x \leq y)$   $\mathcal{P} \models \alpha[1]$  but  $\mathcal{N} \not\models \alpha[1]$  but  $\mathcal{N} \models \alpha[h(1)]$

**Definition 2.20. Automorphism:** An automorphism is an isomorphism of the form  $h : A \rightarrow A$  from  $\mathcal{A}$  onto  $\mathcal{A}$

**Note:** Every structure has a trivial automorphism  $id : A \rightarrow A$

**Definition 2.21. Rigid:** If the only automorphism on  $\mathcal{A}$  is the trivial automorphism, then  $\mathcal{A}$  is called rigid.

**Example 2.18. :** If every element is definable then the structure is rigid. For example  $(\mathbb{N}, 0, S)$  and  $(\mathbb{N}, <)$  every element is definable, therefore the structures are rigid.

**Corollary 2.3.** Let  $h$  be autom of  $\mathcal{A}$ ,  $R \subseteq A^n$  definable in  $\mathcal{A}$  then  $\forall a \in A^n a \in R$  iff  $(h(a_1), \dots, h(a_n)) \in R$  Suppose  $\varphi$  defines  $R$  in  $\mathcal{A}$  we want  $\mathcal{A} \models \varphi[a]$  iff  $\mathcal{A} \models \varphi[h(a_1), \dots, h(a_n)]$  which is true by the homom. thm.

**Note:** Corol can be used to show that some  $R \subseteq A^n$  is not definable in  $\mathcal{A}$

**Example 2.19. :**  $\mathcal{R} = (\mathbb{R}, <)$  then  $\mathbb{N}$  is not definable in  $\mathcal{R}$ . What do automorphisms of  $\mathcal{R}$  look right?  $h : \mathbb{R} \rightarrow \mathbb{R}$  is a bijection and  $x < y$  iff  $h(x) < h(y)$  so  $h$  is strictly increasing. for example  $x \mapsto x + \frac{1}{2}$  or  $x \mapsto x^3$ .

## 2.6 UNIQUE READABILITY FOR TERMS

**Definition 2.22. :** We define  $K$  on symbols from which terms are built up (variables, constants, function symbols).  $K(s) = 1 - n$  where  $s$  is a symbol and  $n$  is the number of terms that need to follow  $s$  in order to obtain a term.  $K(x) = 1 = K(c)$  and  $K(f) = 1 - n$  where  $f$  is an  $n$ -ary function symbol We now extend  $K$  to the set of all expressions which are built up from above symbols (variables, constants, function symbols):  $K(s_1, \dots, s_n) = K(s_1) + \dots + K(s_n)$  (unique because no symbol is a finite sequence of other symbols)

**Lemma 2.1. :**  $t$  a term then  $K(t) = 1$

*Proof.*  $K(x) = 1 = K(c)$  and  $K(ft_1, \dots, t_n) = 1 - n + n = 1$   $\square$

**Definition 2.23. :** A terminal segment of string of symbols  $(s_1, \dots, s_n)$  is  $(s_k, s_{k+1}, \dots, s_n)$  for some  $1 \leq k \leq n$ .

**Lemma 2.2. :** Any terminal segment of terms is a concatenation of one or more terms.

*Proof.* True for variables and constants.  $ft_1 \dots t_n$  the only non trivial case is  $t'_k t_{k+1} \dots t_m$  where  $t_k$  is  $t''_k t'_k$   $\square$

**Corollary 2.3.** If  $t_1$  is a proper initial segment of a term  $t$  then its  $K(t_1) < 1$ . proof: let  $t$  be  $t_1 t_2$  where  $t_1$  is a proper initial segment then  $K(t) = 1$  and  $K(t_2) \geq 1$  therefore  $K(t_1) \leq 0$

### Unique readability for terms

The set of terms is freely generated from the set of variables (Var), the set of constant symbols (Const) by the term building operations  $\mathcal{F}_f$  for the function symbols  $f$ .

*Proof.* • disjointment of ranges: Let  $f$  and  $g$  be two distinct function symbols then  $\text{ran } \mathcal{F}_f \cap \text{ran } \mathcal{F}_g = \emptyset$   $\text{ran } \mathcal{F}_f \cap \text{Var} = \emptyset$   $\text{ran } \mathcal{F}_f \cap \text{Const} = \emptyset$

- $\mathcal{F}_f|_{\text{terms}}$  are 1-1: assume  $f t_1 \dots t_n = f t'_1 \dots t'_n$  and assume  $t_1 \neq t'_1$  then one is an initial segment of the other. Then its  $K$ -value has to be less than 1 so it is not a term.  $t_1 = t'_1 \dots t_n = t'_n$ .

□

**Definition 2.24. :** Extend  $K$  as follows:  $K(()) = -1$   $K(()) = 1$   $K(\forall) = 1$   $K(\neg) = 0$   $K(\rightarrow) = -1$   $K(P) = 1 - n$  for an  $n$ -ary rel. symb.  $P$ .  $K(=) = -1$ . Extend  $K$  to the set of all expressions by  $K(s_1, \dots, s_n) = K(s_1) + K(s_n)$  The idea is that  $K$  tells us the number of symbols that at least need to follow to obtain a formula.

**Lemma 2.3. :** for every formula  $\varphi$  it is  $K(\varphi) = 1$

*Proof.* induction on  $\varphi$

□

**Lemma 2.4. :** for every proper initial segment  $\alpha'$  of a fla.  $\alpha$  we have  $K(\alpha') < 1$

**Corollary 2.3.** No proper initial segment of a fla. is a fla.

The set of flas. is freely generated from the set of atomic flas. by operations  $\mathcal{E}_{\neg}, \mathcal{E}_{\rightarrow}, Q_i$

*Proof.* •  $\mathcal{E}_{\neg}, Q_i$  are one to one

- $\mathcal{E}_{\rightarrow}|_{\text{Flas.}}$  then itemwise and use of prev. lemmas
- p.w. disjointness of ranges

□

## 2.7 A PARSING ALGORITHM FOR FIRST ORDER LOGIC

### 2.8 DEDUCTIONS (FORMAL PROOFS)

**Definition 2.25. Modus Ponens:** Rule of inference:

$$\frac{\alpha, \alpha \rightarrow \beta}{\beta}$$

**Definition 2.26. :** A formal proof (deduction) of a fla  $\varphi$  from a set of formulas  $\Sigma$  is a finite sequence of formulas  $(\alpha_0, \alpha_1, \dots, \alpha_n)$  such that  $\alpha_n = \varphi$  and for every  $i < n$   $\alpha_i$  is either a logical axiom or  $\alpha_i \in \Sigma$  or  $\alpha_i$  is obtained from  $\alpha_k$  and  $\alpha_l$  where  $0 \leq k, l < i$  by the use of MP, in particular  $\alpha_k = \beta \rightarrow \alpha_i$  and  $\alpha_l = \beta$ . If a deduction of  $\varphi$  from  $\Sigma$  exists, we say "  $\varphi$  is deducible from  $\Sigma$ " or "  $\varphi$  is a theorem of  $\Sigma$ ".

**Note:** Deductions are not unique. However we do have an induction principle: If a set of formulas contains all logical axioms and all of  $\Sigma$  and is closed under MP, then it contains all theorems of  $\Sigma$ .

**Logical axioms**

**Definition 2.27. Generalization:**  $\psi$  is a generalization of  $\varphi$  if  $\psi = \forall x_{i_1} \dots \forall x_{i_k} \varphi$

**Definition 2.28. Logical axioms:** Let  $x, y$  be variables and  $\alpha, \beta$  formulas. then the logical axioms are generalizations of the following formulas:

1. tautologies
2.  $\forall x \alpha \rightarrow \alpha_t^x$  where  $t$  is substitutable for  $x$  in  $\alpha$
3.  $\forall x(\alpha \rightarrow \beta) \rightarrow (\forall x \alpha \rightarrow \forall x \beta)$
4.  $\alpha \rightarrow \forall x \alpha$  where  $x$  does not occur free in  $\alpha$

if our language contains  $=$  then

1.  $x = x$
2.  $x = y \rightarrow (\alpha \rightarrow \alpha')$  where  $\alpha'$  is obtained from  $\alpha$  by replacing some of the occurrences of  $x$  with  $y$ .

Ad 2) Substitution:

**Definition 2.29. :** Let  $\alpha, \beta$  be formulas,  $x$  a variable and  $t$  a term then  $\alpha_t^x$  is expression obtained from  $\alpha$  by substituting  $t$  for  $x$  We define substitution inductive as follows:

1. if  $\alpha$  is atomic then  $\alpha =$  expression obtained from  $\alpha$  by replacing all  $x$ 's by  $t$ 's
2.  $(\neg \alpha)_t^x = \neg(\alpha_t^x)$
3.  $(\alpha \rightarrow \beta)_t^x = (\alpha_t^x \rightarrow \beta_t^x)$
4.  $(\forall y \alpha)_t^x = \begin{cases} \forall y(\alpha_t^x) & \text{iff } x \neq y \\ \forall x \alpha & \text{iff } x = y \end{cases}$

**Example 2.20. :**

- $\alpha_x^x = \alpha$
- Let  $\alpha = \neg \forall y x = y$  what is  $\forall x \alpha \rightarrow \alpha_z^x$ ?

$$\forall x \neg \forall y x = y \rightsquigarrow \neg \forall y z = y$$

What is  $\forall x \alpha \rightarrow \alpha_y^x$   $\forall x \neg \forall y x = y$  is true in all structures with a universe  $A$  with  $|A| \geq 2$ .

$$\forall x \neg \forall y x = y \rightsquigarrow \neg \forall y y = y$$

and  $\neg \forall y y = y$  is an antitautology (it is always false).

•

So we have to define substitutable

**Definition 2.30. :** Let  $x$  be a variable,  $t$  a term. Then  $t$  is substitutable for  $x$  in  $\alpha$  if

1.  $\alpha$  atomic then  $t$  is SA for  $x$  in  $\alpha$
2. then  $t$  is SA for  $x$  in  $\neg \alpha$  iff then  $t$  is SA for  $x$  in  $\alpha$
3. then  $t$  is SA for  $x$  in  $\alpha \rightarrow \beta$  iff then  $t$  is SA for  $x$  in  $\alpha$  and  $\beta$
4. then  $t$  is SA for  $x$  in  $\forall y \alpha$  iff either
  - $x$  does not occur free in  $\forall y \alpha$  or
  - $y$  does not occur in  $t$  and  $t$  is SA for  $x$  in  $\alpha$

**Example 2.21. :** For instance the following is a logical axiom.

$$\forall x_3(\forall x_1(Ax_1 \rightarrow \forall x_2 Ax_2) \rightarrow (Ax_2 \rightarrow \forall x_2 Ax_2))$$

It is a generalization of  $\forall x_1(Ax_1 \rightarrow \forall x_2 Ax_2) \rightarrow (Ax_2 \rightarrow \forall x_2 Ax_2)$  which is by point two a substitution with  $\alpha = Ax_1 \rightarrow \forall x_2 Ax_2$ . Then  $\alpha_{x_2}^{x_1} = Ax_2 \rightarrow \forall x_2 Ax_2$  And  $x_2$  is indeed substitutable for  $x_1$  in  $\alpha$  because it does not get bounded.

$$\forall x_1(\forall x_2 Bx_1 x_2 \rightarrow \forall x_2 Bx_2 x_2)$$

is a generalization of point (2), but  $x_2$  is not substitutable for  $x_1$  in  $\alpha =$ , therefore it is not a logical axiom.

**Ad (1): tautologies**

**Definition 2.31. :** Tautologies are the formulas obtained from tautologies of propositional logic by replacing all propositional atoms by formulas of first order logic.

An alternative definition is: Divide all formulas of first order logic into two groups:

1. atomic formulas and generalizations of first order formulas (these are called prime formulas)
2. all other formulas i.e. of the form  $\neg\alpha$  and  $\alpha \rightarrow \beta$  (non-prime formulas)

So any first order formula is built up from the prime formulas using finitely many times the formula building operations.  $\mathcal{E}_\neg \mathcal{E}_\rightarrow$  We have unique readability because the set of formulas is freely generated.

**Example 2.22. :**

$$\neg(\forall y(Px \rightarrow Py)) \rightarrow (Px \rightarrow \forall y\neg Py)$$

is built up from  $\neg(\forall y(Px \rightarrow Py))$  and  $Px \rightarrow \forall y\neg Py$ . which itself  $\forall y(Px \rightarrow Py)$  and  $Px$  and  $\forall y\neg Py$  where they are all prime formulas.

**Example 2.23. :** Is the following a tautology?

$$(\forall y(\neg Py) \rightarrow \neg Px) \rightarrow (Px \rightarrow \neg \forall y\neg Py)$$

We construct the construction tree into prime formulas and then assign truth values to them and evaluate the truth value of the whole formula. It is indeed a tautology.

**Note:**

- $\forall x(Px \rightarrow Px)$  is a prime formula which corresponds to a propositional atom, and therefore not a tautology. But it is a generalization of a tautology and therefore by (1) a logical axiom.
- $\forall xPx \rightarrow Px$  is not a tautology but is a logical axiom by group (2).

**Note:**  $\Gamma \models_{\text{taut}} \varphi$  from propositional logic can be translated to first order logic.

**Lemma 2.5. :** If  $\Gamma \models_{\text{taut}} \varphi$  then  $\Gamma \models \varphi$

*Proof.* Problem set. □

**Note:** The converse fails. For instance  $\forall xPx \models Pc$ . However  $Pc$  is a different propositional atom then  $\forall xPx$  they have no connection between them when viewed in propositional logic.

We will prove  $\Gamma \models \varphi$  iff  $\Gamma \vdash \varphi$  (the first direction is completeness and the converse soundness.)

**Theorem 2.4. :**  $\Gamma \vdash \varphi$  iff  $\Gamma \cup \Lambda \models_{\text{taut}} \varphi$

*Proof.* • Let  $\Gamma \vdash \varphi$  and  $v$  be a truth assignment that satisfies every element in  $\Gamma \cup \Lambda$ . Induction on deduction of  $\varphi$  from  $\Gamma$ .

- if  $\varphi \in \Gamma \cup \Lambda$  then we are done
- if  $\varphi$  is obtained from  $\alpha, \alpha \rightarrow \varphi$  by MP then  $v$  satisfies  $\alpha$  and  $\alpha \rightarrow \varphi$   
 $\{\alpha, \alpha \rightarrow \varphi\} \models_{\text{taut}} \varphi$

- Assume  $\Gamma \cup \Lambda \models_{\text{taut}} \varphi$ . Then by the compactness theorem for propositional logic there are  $\gamma_1, \dots, \gamma_n \in \Gamma$  and  $\lambda_1, \dots, \lambda_m \in \Lambda$  such that

$$\gamma_1 \rightarrow \gamma_2 \rightarrow \dots \rightarrow \gamma_n \rightarrow \lambda_1 \rightarrow \dots \rightarrow \lambda_m$$

is a tautology (always grouped to the left) because  $\Gamma \cup \{\alpha\} \models_{\text{taut}} \beta$  iff  $\Gamma \models_{\text{taut}} (\alpha \rightarrow \beta)$  □

## 2.9 GENERALIZATION AND DEDUCTION THEOREM

TODO evt noch sections

## CHAPTER 3

# Boolean Algebra

**Definition 3.1. Boolean Algebra:** A boolean algebra is a set  $B$  with

- distinguished elements  $0, 1$  (called zero and unit of  $B$ )
- a unary operation  $'$  on  $B$  (called **complementation** )
- two binary operations  $\vee$  called **join** and  $\wedge$  called **meet** s.t. for all  $x, y, z \in B$

1.  $x \vee 0 = x$        $x \wedge 1 = x$
2.  $x \vee x' = 1$        $x \wedge x' = 0$
3.  $x \vee y = y \vee x$        $x \wedge y = y \wedge x$
4.  $(x \vee y) \vee z = x \vee (y \vee z)$        $(x \wedge y) \wedge z = x \wedge (y \wedge z)$
5.  $x \vee (y \wedge z) = (x \vee y) \wedge (x \vee z)$        $x \wedge (y \vee z) = (x \wedge y) \vee (x \wedge z)$

**Example 3.1. :** Let  $S$  be a set,  $B := \mathcal{P}(S)$  the power set of  $S$ ,  $0 := \emptyset$  and  $1 := S$ ,

$$' : \mathcal{P}(S) \rightarrow \mathcal{P}(S), x' := S \setminus x \quad x \vee y := x \cup y, \quad x \wedge y := x \cap y \text{ for } x, y \in \mathcal{P}(S)$$

**Lemma 3.1. :** Let  $(B, ', \vee, \wedge, 0, 1)$  be a boolean algebra. Then it holds

- a)  $0' = 1, 1' = 0$
- b)  $x \vee x = x, x \wedge x = x$
- c)  $(x')' = x$
- d)  $(x \vee y)' = x' \wedge y', (x \wedge y)' = x' \vee y'$
- e)  $x \vee y = y$  iff  $x \wedge y = x$

**Lemma 3.2. :**

- a)  $x \leq y :\Leftrightarrow x \vee y = y$  defines a partial ordering on  $B$  (inclusion) and it holds
- b)  $x \vee y$  is the least upper bound of  $\{x, y\}$  in  $B$   
 $x \wedge y$  is the greatest lower bound of  $\{x, y\}$  in  $B$
- c)  $0 \leq x \leq 1$  for all  $x \in B$

**Note:** A boolean algebra is a complemented distributive lattice.

**Definition 3.2. Opposite of boolean algebra:** Let  $(B, ', \vee, \wedge, 0, 1)$  be a boolean algebra. The boolean algebra  $B^{\text{op}}$  is defined by

$$B^{\text{op}} := B, \quad 0^{\text{op}} := 1, \quad 1^{\text{op}} := 0, \quad ' \text{ stays the same as for } B, \quad \vee^{\text{op}} := \wedge, \quad \wedge^{\text{op}} := \vee$$

Note:  $(B^{\text{op}})^{\text{op}} = B$

**Definition 3.3. Subalgebra:** A subalgebra of  $B$  is a subset  $A \subseteq B$  s.t.  $0, 1 \in A$  and  $A$  is closed under  $', \wedge, \vee$ . The subalgebra generated by  $P \subseteq B$  is defined to be the smallest subalgebra containing  $P$ . Equivalently it is the intersection of all Subalgebras of  $B$  that contain  $P$ .

**Example 3.2. Power set algebra:** Let  $S$  be a set then  $\mathcal{P}(S)$  defines a boolean algebra on  $S$ .  $B := \{x \in \mathcal{P}(S) : x \text{ is finite or cofinite}\}$  is a subalgebra of  $\mathcal{P}(S)$  w/ set of generators  $\{\{s\} : s \in S\}$

**Note:** We will prove the Tarski-Stone Theorem: every boolean algebra is isomorphic to an algebra on a set.

**Example 3.3. Lindenbaum Algebra of  $\Sigma$ :** Let  $A$  be a set of prop. atoms,  $\text{Prop}(A)$  the set of prop. generated by  $A$ . Further let  $\Sigma \subseteq \text{Prop}(A)$  and  $p, q, r$  range over  $\text{Prop}(A)$ . We say  $p$  is  $\Sigma$ -equivalent to  $q$  iff  $\Sigma \models_{\text{taut}} p \leftrightarrow q$ .  $\Sigma$ -Equivalence is an equivalent relation on  $\text{Prop}(A)$  and  $\text{Prop}(A)/\Sigma$  is a boolean algebra with

$$0 := \perp/\Sigma, \quad 1 := \top/\Sigma, \quad (p/\Sigma)' := (\neg p)/\Sigma, \quad (p/\Sigma \vee q/\Sigma) := (p \vee q)/\Sigma, \quad (p/\Sigma \wedge q/\Sigma) := (p \wedge q)/\Sigma$$

a set of generators is  $\{a/\Sigma : a \in A\}$

**Definition 3.4. Homomorphisms of boolean algebras:** Let  $B, C$  be boolean algebras. A map  $\phi : B \rightarrow C$  is a (homo)morphism of boolean algebras iff  $\forall x, y \in B$  it holds

- $\phi(0_B) = 0_C$
- $\phi(x') = \phi(x)'$
- $\phi(x \vee y) = \phi(x) \vee \phi(y)$
- $\phi(x \wedge y) = \phi(x) \wedge \phi(y)$

If  $\phi : B \rightarrow C$  is bijective too, we call  $\phi$  an isomorphism and  $\phi^{-1} : C \rightarrow B$  is also a morphism of boolean algebras.

**Note:**  $\phi(B)$  is subalgebra of  $C$

**Example 3.4. :** Let  $S, T$  be sets then a function  $f : S \rightarrow T$  induces a morphism of boolean algebras  $\mathcal{P}(T) \rightarrow \mathcal{P}(S) : y \mapsto f^{-1}(y)$ . If  $S \subseteq T$  and  $f$  the inclusion map  $S \hookrightarrow T$  then we get a boolean algebra morphism  $Y \rightarrow Y \cap S$ .

•  $\text{id}_B : B \rightarrow B$  •  $x \mapsto x' : B \rightarrow B^{\text{op}}$  are both isomorphism

**Note:** A boolean algebra morphism  $\phi : B \rightarrow C$  is injective iff  $\ker \phi = 0_B$

**Lemma 3.3. :** Let  $X_1, \dots, X_m \subseteq S$  and  $\mathcal{A}$  a boolean algebra on  $S$  generated by  $\{X_1, \dots, X_m\}$ . Then  $\mathcal{A}$  is finite and isomorphic to  $\mathcal{P}(\{1, 2, \dots, n\})$  for some  $n \leq 2^m$ .

*Proof.* TODO □

**Definition 3.5. Trivial algebras:**

- $B$  is trivial if  $|B| = 1$  (equivalently  $0 = 1 \in B$ ) according to Lemma 3.3  $B$  is isomorphic to  $\mathcal{P}(\emptyset)$
- If  $|S| = 1$  then  $|\mathcal{P}(S)| = 2$  TODO

**Definition 3.6. Ideal:** An ideal of  $B$  is a subset of  $I \subseteq B$  s.t.

$$(I1) \quad 0 \in I$$

$$(I2) \quad \forall a, b \in B \text{ it holds} \quad a \leq b \text{ and } b \in I \implies a \in I \quad \text{and} \quad a, b \in I \implies a \vee b \in I$$

**Example 3.5. :**  $F_{\text{in}} = \{F \subseteq S : F \text{ finite}\}$  is ideal in  $\mathcal{P}(S)$ .

**Note:** If  $I$  is an ideal of  $B$  then  $I \vee b := \{x \in B : x = a \vee b \text{ for some } a \in I\}$  is the smallest ideal w/ respect of  $\subseteq$  of  $B$  that contains  $I \cup \{b\}$ .

**Example 3.6. :**

- For a boolean algebra morphism  $\phi : B \rightarrow C$  the kernel  $\ker(\phi)$  is an ideal in  $B$ .
- If  $I$  is an ideal in  $B$  then  $a =_I b \iff a \vee x = b \vee x \text{ for some } x \in I$  defines an equivalent relation and  $B/_I$  is a boolean algebra w/

$$0 := 0/_I \quad 1 := 1/_I \quad (a/_I)' := a'/_I \quad a/_I \vee b/_I := (a \vee b)/_I \quad a/_I \wedge b/_I := (a \wedge b)/_I$$

Then  $\phi : B \rightarrow B/_I : b \mapsto b/_I$  is a boolean algebra morphism w/  $\ker(\phi) = I$

## CHAPTER 4

# Set Theory

**Example 4.1. Russel's paradox:** Let  $A = \{a : a \notin a\}$ . If any collection of elements is a set, then  $A$  would be a set. Question: is  $A \in A$ ? if yes, then  $A \notin A$ , if not then  $A \in A$

Trying to resolve this, we will introduce the ZFC (Zermelo-Frankel axioms w/ choice) System. Let  $\mathcal{L} = \{\in\}$  be a Language of first order, where  $\in \dots$  binary relation "being element of". For  $(\mathcal{U}, \in)$  If  $(\mathcal{U}, \in) \models \text{ZFC}$ , then the elements of the universe  $\mathcal{U}$  are called sets.

TODO

### 4.1 AXIOMS OF ZFC

**Definition 4.1. Axiom of extensionality:**

$$\forall x \forall y (x = y \leftrightarrow \forall u (u \in x \leftrightarrow u \in y))$$

**Definition 4.2. Pairing Axiom:** for any two sets  $a, b$  one can form a set whose elements are precisely  $a, b$

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

Our notation will be  $z = \{x, y\}$

**Note:**  $\{x, y\}$  is unique by [Definition 4.1](#)

**Lemma 4.1. :** Let  $x, y$  be sets. We define  $(x, y) := \{\{x\}, \{x, y\}\}$ . Then it holds  $(x, y) = (a, b)$  iff  $x = a$  and  $y = b$

*Proof.* • if  $x = y$ , then  $(x, y) = \{\{x\}\}$  therefore  $a = b$  and by [Definition 4.1](#) it holds  $x = a$ .

- if  $x \neq y$ , then  $\{\{x\}, \{x, y\}\} = \{\{a\}, \{a, b\}\}$  iff  $\{x\} = \{a\}$  and  $\{x, y\} = \{a, b\}$ . That is, iff  $x = a$  and  $y = b$ .

□

TODO ordered n-tuples

**Definition 4.3. Union Axiom:** For every set  $x$  there is a set  $z$  consisting of all elements of the elements of  $x$ .

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

We call  $z$  the union of  $x$ , notation:  $\bigcup_x := z$

**Definition 4.4. Power set Axiom:** Let  $x \subseteq y$  be the abbreviation for  $\forall z (z \in x \rightarrow z \in y)$ . The **Powerset Axiom** states, that for every set  $x$  there exists a set  $z$  consisting of all subsets  $y \subseteq x$  that are themselves sets.

$$\forall x \exists z \forall y (y \in z \leftrightarrow y \subseteq x)$$

Notation:  $\mathcal{P}(x) := z$ .

TODO class relations

**Definition 4.5. Axiom of replacement / substitution:** Let  $\varphi(x, y, \underline{a})$  a  $\mathcal{L}$ -f.a., w/ free variables among  $x, y$  and set-parameters  $\underline{a}$ . Suppose  $\varphi$  defines a class function on  $\mathcal{U}$ , then the following is an axiom:

$$\forall u \exists z \forall y (y \in z \leftrightarrow \exists x (x \in u \wedge \varphi(x, y, \underline{a})))$$

i.e. the image of a set under a class function is a set.

**Definition 4.6. Axiom scheme of comprehension:** TODO



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