# CATEGORY THEORY AND LAMBDA CALCULUS

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# CONTENTS

## ABSTRACT

This is the abstract. It should not be written until the end.

Part I

# CATEGORY THEORY

# 1

# CATEGORIES

## 1.1 DEFINITION OF CATEGORY

**Definition 1.** A **category** $\mathcal{C}$, as defined in [Lan78], is given by

- $\mathcal{C}_0$, a collection[1] whose elements are called **objets**, and
- $\mathcal{C}_1$, a collection whose elements are called **morphisms**.

Every morphism $f \in \mathcal{C}_1$ has two objects assigned: a **domain**, written as $\operatorname{dom}(f) \in \mathcal{C}_0$, and a **codominio**, written as $\operatorname{cod}(f) \in \mathcal{C}_0$; a common notation for such morphism is

$$f \colon \operatorname{dom}(f) \to \operatorname{cod}(f).$$

Given two morphisms $f \colon A \to B$ and $g \colon B \to C$ there exists a **composition morphism**, written as $g \circ f \colon A \to C$. Morphism composition is a binary associative operation with identity elements $\mathrm{id}_A \colon A \to A$, that is

$$h \circ (g \circ f) = (h \circ g) \circ f \quad \text{and} \quad f \circ \mathrm{id}_A = f = \mathrm{id}_B \circ f.$$

1 : We use the term *collection* to denote some unspecified formal notion of compilation of "things" that could be given by sets or proper classes. We will want to define categories whose objects are all the possible sets and we will need the objects to form a proper class.

Part II

# LAMBDA CALCULUS

# 2

# UNTYPED $\lambda$-CALCULUS

The $\lambda$-**calculus** is a collection of formal systems, all of them based on the lambda notation discovered by Alonzo Church in the 1930s while trying to develop a foundational notion of function on mathematics.

The **untyped** or **pure lambda calculus** is, syntactically, the simplest of those formal systems. This presentation of the untyped lambda calculus will follow [HS08] and [Sel13].

## 2.1 DEFINITION

**Definition 2.** The $\lambda$-**terms** are defined inductively as

- every *variable*, taken from an infinite and numerable set $\mathcal{V}$ of variables, and usually written as lowercase single letters (x,y,z,. . . ), is a $\lambda$-term.
- given two $\lambda$-terms $M, N$; its *application*, $MN$ is a $\lambda$-term.
- given a $\lambda$-term $M$ and a variable $x$, its *abstraction*, $\lambda x.M$ is a lambda term.

They can be also defined by the following BNF

$$\mathtt{Exp} ::= x \mid (\mathtt{Exp}\ \mathtt{Exp}) \mid (\lambda x.\mathtt{Exp})$$

where $x \in \mathcal{V}$ is any variable.

By convention, we omit outermost parentheses and assume left-associativity, i.e., $MNP$ will mean $(MN)P$. Multiple $\lambda$-abstractions can be also contracted to a single multivariate abstraction; thus $\lambda x.\lambda y.M$ can become $\lambda x, y.M$.

## 2.2 FREE AND BOUND VARIABLES, SUBSTITUTION

Any ocurrence of a variable $x$ inside the *scope* of a lambda is said to be bound; and any not bound variable is said to be free. We can define formally the set of free variables as follows.

**Definition 3.** The **set of free variables** of a term $M$ is defined inductively as

$$\begin{aligned} FV(x) &= \{x\}, \\ FV(MN) &= FV(M) \cup FV(N), \\ FV(\lambda x.M) &= FV(M) \setminus \{x\}. \end{aligned}$$

A free ocurrence of a variable can be substituted by a term. This should be done avoiding the unintended bounding of free variables which happens when a variable is substituted inside of the scope of a binder with the same name, as in the following example, where we substitute $y$ by $(\lambda z.xz)$ on $(\lambda x.yx)$ and the second free variable $x$ gets bounded by the first binder

$$(\lambda x.yx) \overset{y \mapsto (\lambda z.xz)}{\longrightarrow} (\lambda x.(\lambda z.xz)x).$$

To avoid this, the $x$ should be renamed before the substitution.

**Definition 4.** The **substitution** of a variable $x$ by a term $N$ on $M$ is defined inductively as

$$\begin{aligned} x[N/x] &\equiv N, & \\ y[N/x] &\equiv y, & \\ (MP)[N/x] &\equiv (M[N/x])(P[N/x]), & \\ (\lambda x.P)[N/x] &\equiv \lambda x.P, & \\ (\lambda y.P)[N/x] &\equiv \lambda y.P[N/x] & \text{if } y \notin FV(N), \\ (\lambda y.P)[N/x] &\equiv \lambda z.P[z/y][N/x] & \text{if } y \in FV(N); \end{aligned}$$

where, in the last clause, $z$ is a fresh unused variable.

We could define a criterion for choosing exactly what this new variable should be, or simply accept that our definition will not be well-defined, but well-defined up to a change on the name of the variables. This equivalence relation will be defined formally on the next section. In practice, it is common to follow the *Barendregt's variable convention* which simply assumes that bound variables have been renamed to be distinct.

## 2.3 α-EQUIVALENCE

**Definition 5.** **$\alpha$-equivalence** is the smallest relation $=_\alpha$ on $\lambda$-terms which is an equivalence relation, i.e.,

- it is *reflexive*, $M =_\alpha M$;
- it is *symmetric*, if $M =_\alpha N$, then $N =_\alpha M$;
- and it is *transitive*, if $M =_\alpha N$ and $N =_\alpha P$, then $M =_\alpha P$;

and it is compatible with the structure of lambda terms,

- if $M =_\alpha M'$ and $N =_\alpha N'$, then $MN =_\alpha M'N'$;
- if $M =_\alpha M'$, then $\lambda x.M =_\alpha \lambda x.M'$;
- if $y$ does not appear on $M$, $\lambda x.M =_\alpha \lambda y.M[y/x]$.

$\alpha$-equivalence formally captures the fact that the name of a bound variable can be changed without changing the properties of the term. This idea appears recurrently on mathematics; the renaming of the variable of integration is an example of $\alpha$-equivalence.

$$\int_0^1 x^2 \, dx = \int_0^1 y^2 \, dy$$

## 2.4 $\beta$-REDUCTION

The core idea of evaluation in $\lambda$-calculus is captured by the notion of $\beta$-reduction.
**Definition 6.** The **single-step $\beta$-reduction** is the smallest relation on $\lambda$-terms capturing the notion of evaluation

$$(\lambda x.M)N \to_\beta M[N/x],$$

and some congruence rules that preserve the structure of $\lambda$-terms, such as

- $M \to_\beta M'$ implies $MN \to_\beta M'N$ and $MN \to_\beta MN'$;
- $M \to_\beta M'$ implies $\lambda x.M \to_\beta \lambda x.M'$.

The reflexive transitive closure of $\to_\beta$ is written as $\twoheadrightarrow_\beta$. The symmetric closure of $\twoheadrightarrow_\beta$ is called **$\beta$-equivalence** and written as $=_\beta$ or simply $=$.

## 2.5 $\eta$-REDUCTION

The idea of function extensionality in $\lambda$-calculus is captured by the notion of $\eta$-reduction. Function extensionality implies the equality of any two terms that define the same function over any argument.
**Definition 7.** The $\eta$-reduction is the smallest relation on $\lambda$-terms satisfiying the same congruence rules as $\beta$-reduction and the following axiom

$$\lambda x.Mx \to_\eta M, \text{ for any } x \notin \mathrm{FV}(M).$$

We define single-step $\beta\eta$-reduction as the union of $\beta$-reduction and $\eta$-reduction. This will be written as $\to_{\beta\eta}$, and its reflexive transitive closure will be $\twoheadrightarrow_{\beta\eta}$.

## 2.6 CONFLUENCE

**Definition 8.** A relation $\to$ is **confluent** if, given its reflexive transitive closure $\twoheadrightarrow$, $M \twoheadrightarrow N$ and $M \twoheadrightarrow P$ imply the existence of some $Z$ such that $N \twoheadrightarrow Z$ and $P \twoheadrightarrow Z$.

Given any binary relation $\to$ of which $\twoheadrightarrow$ is its reflexive transitive closure, we can consider three seemingly related properties

- the **confluence** or Church-Rosser property we have just defined.
- the **quasidiamond property**, which assumes $M \to N$ and $M \to P$.
- the **diamond property**, which is defined substituting $\twoheadrightarrow$ by $\to$ on the definition on confluence.

Diagrammatically, the three properties can be represented as

$$\begin{tikzcd}
 & M \arrow[dl, two heads] \arrow[dr, two heads] & \\
N \arrow[dr, dashed, two heads] & & P \arrow[dl, dashed, two heads] \\
 & Z &
\end{tikzcd}
\qquad
\begin{tikzcd}
 & M \arrow[dl] \arrow[dr] & \\
N \arrow[dr, dashed, two heads] & & P \arrow[dl, dashed, two heads] \\
 & Z &
\end{tikzcd}
\qquad
\begin{tikzcd}
 & M \arrow[dl] \arrow[dr] & \\
N \arrow[dr, dashed] & & P \arrow[dl, dashed] \\
 & Z &
\end{tikzcd}$$

and the implication relation between them is that the diamond relation implies confluence; while the quasidiamond does not. Both claims are easy to prove, and they show us that, in order to prove confluence for a given relation, we need to prove the diamond property instead of try to prove it from the quasidiamond property, as a naive attempt of proof would try.

The statement of $\twoheadrightarrow_\beta$ and $\twoheadrightarrow_{\beta\eta}$ being confluent is what we are going to call the Church-Rosser theorem. The definition of a relation satisfying the diamond property and whose reflexive transitive closure is $\twoheadrightarrow_{\beta\eta}$ will be the core of our proof.

## 2.7 THE CHURCH-ROSSER THEOREM

The proof presented here is due to Tait and Per Martin-Löf; an earlier but more convoluted proof was discovered by Alonzo Church and Barkley Rosser in 1935. It is based on the idea of parallel one-step reduction.

**Definition 9** (Parallel one-step reduction)**.** We define the **parallel one-step reduction** relation, $\rhd$ as the smallest relation satisfying that, assuming $P \rhd P'$ and $N \rhd N'$, the following properties of

- reflexivity, $x \rhd x$;

- parallel application, $PN \rhd P'N'$;
- congruence to $\lambda$-abstraction, $\lambda x.N \rhd \lambda x.N'$;
- parallel substitution, $(\lambda x.P)N \rhd P'[N'/x]$;
- and extensionality, $\lambda x.Px \rhd P'$, if $x \notin \mathrm{FV}(P)$,

hold.

Using the first three rules, it is trivial to show that this relation is in fact reflexive.
**Lemma 1.** *The reflexive transitive closure of $\rhd$ is $\twoheadrightarrow_{\beta\eta}$. In particular, given any $M, M'$,*

1. *if $M \rightarrow_{\beta\eta} M'$, then $M \rhd M'$.*
2. *if $M \rhd M'$, then $M \twoheadrightarrow_{\beta\eta} M'$;*

*Proof.* 1. We can prove this by exhaustion and structural induction on $\lambda$-terms, the possible ways in which we arrive at $M \rightarrow M'$ are

   - $(\lambda x.M)N \rightarrow M[N/x]$; where we know that, by parallel substitution and reflexivity $(\lambda x.M)N \rhd M[N/x]$.
   - $MN \rightarrow M'N$ and $NM \rightarrow NM'$; where we know that, by induction $M \rhd M'$, and by parallel application and reflexivity, $MN \rhd M'N$ and $NM \rhd NM'$.
   - congruence to $\lambda$-abstraction, which is a shared property between the two relations where we can apply structural induction again.
   - $\lambda x.Px \rightarrow P$, where $x \notin \mathrm{FV}(P)$ and we can apply extensionality for $\rhd$ and reflexivity.

2. We can prove this by induction on any derivation of $M \rhd M'$. The possible ways in which we arrive at this are

   - the trivial one, reflexivity.
   - parallel application $NP \rhd N'P'$, where, by induction, we have $P \twoheadrightarrow P'$ and $N \twoheadrightarrow N'$. Using two steps, $NP \twoheadrightarrow N'P \twoheadrightarrow N'P'$ we prove $NP \twoheadrightarrow N'P'$.
   - congruence to $\lambda$-abstraction $\lambda x.N \rhd \lambda x.N'$, where, by induction, we know that $N \twoheadrightarrow N'$, so $\lambda x.N \twoheadrightarrow \lambda x.N'$.
   - parallel substitution, $(\lambda x.P)N \rhd P'[N'/x]$, where, by induction, we know that $P \twoheadrightarrow P'$ and $N \twoheadrightarrow N'$. Using multiple steps, $(\lambda x.P)N \twoheadrightarrow (\lambda x.P')N \twoheadrightarrow (\lambda x.P')N' \rightarrow P'[N'/x]$.
   - extensionality, $\lambda x.Px \rhd P'$, where by induction $P \twoheadrightarrow P'$, and trivially, $\lambda x.Px \twoheadrightarrow \lambda x.P'x$.

Because of this, the reflexive transitive closure of $\rhd$ should be a subset and a superset of $\twoheadrightarrow$ at the same time. □

**Lemma 2** (Substitution Lemma)**.** *Assuming $M \rhd M'$ and $U \rhd U'$, $M[U/y] \rhd M'[U'/y]$.*

*Proof.* We apply structural induction on derivations of $M \rhd M'$, depending on what the last rule we used to derive it was.

- Reflexivity, $M = x$. If $x = y$, we simply use $U \rhd U'$; if $x \neq y$, we use reflexivity on $x$ to get $x \rhd x$.

- Parallel application. By induction hypothesis, $P[U/y] \rhd P'[U'/y]$ and $N[U/y] \rhd N'[U'/y]$, hence $(PN)[U/y] \rhd (P'N')[U'/y]$.
- Congruence. By induction, $N[U/y] \rhd N'[U'/y]$ and $\lambda x.N[U/y] \rhd \lambda x.N'[U'/y]$.
- Parallel substitution. By induction, $P[U/y] \rhd P'[U'/y]$ and $N[U/y] \rhd N[U'/y]$, hence $((\lambda x.P)N)[U/y] \rhd P'[U'/y][N'[U'/y]/x] = P'[N'/x][U'/y]$.
- Extensionality, given $x \notin \mathrm{FV}(P)$. By induction, $P \rhd P'$, hence $\lambda x.P[U/y]x \rhd P'[U'/y]$.

Note that we are implicitely assuming the Barendregt's variable convention; all variables have been renamed to avoid clashes. □

**Definition 10** (Maximal parallel one-step reduct)**.** The **maximal parallel one-step reduct** $M^*$ of a $\lambda$-term $M$ is defined inductively as

- $x^* = x$;
- $(PN)^* = P^*N^*$;
- $((\lambda x.P)N)^* = P^*[N^*/x]$;
- $(\lambda x.N)^* = \lambda x.N^*$;
- $(\lambda x.Px)^* = P^*$, given $x \notin \mathrm{FV}(P)$.

**Lemma 3** (Diamond property of parallel reduction)**.** *Given any $M'$ such that $M \rhd M'$, $M' \rhd M^*$. Parallel one-step reduction has the diamond property.*

*Proof.* We apply again structural induction on the derivation of $M \rhd M'$.

- Reflexivity gives us $M' = x = M^*$.
- Parallel application. By induction, we have $P \rhd P^*$ and $N \rhd N^*$; depending on the form of $P$, we have
  - $P$ is not a $\lambda$-abstraction and $P'N' \rhd P^*N^* = (PN)^*$.
  - $P = \lambda x.Q$ and $P \rhd P'$ could be derived using congruence to $\lambda$-abstraction or extensionality. On the first case we know by induction hypothesis that $Q' \rhd Q^*$ and $(\lambda x.Q')N' \rhd Q^*[N^*/x]$. On the second case, we can take $P = \lambda x.Rx$, where, $R \rhd R'$. By induction, $(R'x) \rhd (Rx)^*$ and now we apply the substitution lemma to have $R'N' = (R'x)[N'/x] \rhd (Rx)^*[N^*/x]$.
- Congruence. Given $N \rhd N'$; by induction $N' \rhd N^*$, and depending on the form of $N$ we have two cases
  - $N$ is not of the form $Px$ where $x \notin \mathrm{FV}(P)$; we can apply congruence to $\lambda$-abstraction.
  - $N = Px$ where $x \notin \mathrm{FV}(P)$; and $N \rhd N'$ could be derived by parallel application or parallel substitution. On the first case, given $P \rhd P'$, we know that $P' \rhd P^*$ by induction hypothesis and $\lambda x.P'x \rhd P^*$ by extensionality. On the second case, $N = (\lambda y.Q)x$ and $N' = Q'[x/y]$, where $Q \rhd Q'$. Hence $P \rhd \lambda y.Q'$, and by induction hypothesis, $\lambda y.Q' \rhd P^*$.
- Parallel substitution, with $N \rhd N'$ and $Q \rhd Q'$; we know that $M^* = Q^*[N^*/x]$ and we can apply the substitution lemma to get $M' \rhd M^*$.
- Extensionality. We know that $P \rhd P'$ and $x \notin \mathrm{FV}(P)$, so by induction hypothesis we know that $P' \rhd P^* = M^*$.

□

**Theorem 1** (Church-Rosser Theorem). *The relation $\twoheadrightarrow_{\beta\eta}$ is confluent.*

*Proof.* Parallel reduction, $\triangleright$, satisfies the diamond property, which implies the Church-Rosser property. Its reflexive transitive closure is $\twoheadrightarrow_{\beta\eta}$, whose diamond property implies confluence for $\rightarrow_{\beta\eta}$. □

# 3

# SIMPLY TYPED LAMBDA CALCULUS

We will give now a presentation of the **simply-typed lambda calculus** based on [Sel13].

## 3.1 SIMPLE TYPES

We start assuming that a set of **basic types** exists. Those basic types would correspond, in a programming language interpretation, with things like the type of strings or the type of integers. We will also assume that a **unit** type, 1 exists; the unit type will have only one inhabitant.

**Definition 11.** The set of **simple types** is given by the following Backus-Naur form

$$\mathtt{Type} ::= 1 \mid \iota \mid \mathtt{Type} \to \mathtt{Type} \mid \mathtt{Type} \times \mathtt{Type}$$

where 1 is a one-element type and $\iota$ is any *basic type*.

That is to say that, for every two types $A, B$, there exist a **function type** $A \to B$ and a **pair type** $A \times B$.

## 3.2 RAW TYPED LAMBDA TERMS

We will now define the terms of the typed lambda calculus.

**Definition 12.** The set of **typed lambda terms** is given by the BNF

$$\mathtt{Term} ::= * \mid x \mid \mathtt{TermTerm} \mid \lambda x^{\mathtt{Type}}.\mathtt{Term} \mid \langle \mathtt{Term}, \mathtt{Term} \rangle \mid \pi_1 \mathtt{Term} \mid \pi_2 \mathtt{Term}$$

Besides the previously considered term application and a special element $*$ which will be the unique inhabitant of the type 1; we now introduce a typed lambda abstraction and an explicit construction of the pair element with its projections.

## 3.3 TYPING RULES FOR THE SIMPLY-TYPED LAMBDA CALCULUS

The set of raw typed lambda terms contains some meaningless terms under our type interpretation, such as $\pi_1(\lambda x^A.M)$. **Typing rules** will give them the desired semantics; only a subset of these raw lambda terms will be typeable.

**Definition 13.** A **typing context** is a sequence of typing assumptions $x_1 : A_1, \dots, x_n : A_n$, where no variable appears more than once.

Every typing rule assumes a typing context, usually denoted by $\Gamma$ or by a concatenation of typing contexts written as $\Gamma, \Gamma'$; and a consequence from that context, separated by the $\vdash$ symbol.

1. The type of $*$ is 1, the rule $(*)$ builds this element.

$$(*)\,\frac{}{\Gamma \vdash * : 1}$$

2. The $(var)$ rule simply makes explicit the type of a variable from the context.

$$(var)\,\frac{}{\Gamma, x : A \vdash x : A}$$

3. The $(pair)$ rule allow us to build pairs by their components. It acts as a constructor of pairs.

$$(pair)\,\frac{\Gamma \vdash a : A \qquad \Gamma \vdash b : B}{\Gamma \vdash \langle a, b\rangle : A \times B}$$

4. The $(\pi_1)$ and $(\pi_2)$ rules give the semantics of a product with two projections to the pair terms. If we have a pair $m : A \times B$, then $\pi_1 m : A$ and $\pi_2 m : B$. They act as two different destructors of pairs.

$$(\pi_1)\,\frac{\Gamma \vdash m : A \times B}{\Gamma \vdash \pi_1 m : A} \qquad (\pi_2)\,\frac{\Gamma \vdash m : A \times B}{\Gamma \vdash \pi_2 m : B}$$

5. The $(abs)$ introduces a well-typed lambda abstraction. If we have a $h : B$ term depending on $x : A$, we can create a lambda abstraction from this term. It acts as a constructor of function terms.

$$(abs)\,\frac{\Gamma, x : A \vdash h : B}{\Gamma \vdash \lambda x^A.h : A \to B}$$

6. The $(app)$ rule gives the type of a well-typed application of a lambda term. A term $f : A \to B$ applied to a term $a : A$ is a term of type $B$. It acts as a destructor of function terms.

$$(app)\,\frac{\Gamma \vdash f : A \to B \qquad \Gamma \vdash a : A}{\Gamma \vdash fa : B}$$

**Definition 14.** A term is **typable** if we can assign types to all its variables in such a way that a typing judgment for the type is derivable.

Part III

# MIKROKOSMOS

We have developed **Mikrokosmos**, a lambda calculus interpreter written in the purely functional programming language Haskell [HHJW07]. It aims to provide students with a tool to learn and understand lambda calculus.

# 4 PARSING

## 4.1 MONADIC PARSER COMBINATORS

A common approach to building parsers in functional programming is to model parsers as functions. Higher-order functions on parsers act as *combinators*, which are used to implement complex parsers in a modular way from a set of primitive ones. In this setting, parsers exhibit a monad algebraic structure, which can be used to simplify the combination of parsers. A technical report on **monadic parser combinators** can be found on [HM96].

The use of monads for parsing is discussed firstly in [Wad85], and later in [Wad90] and [HM98]. The parser type is defined as a function taking a `String` and returning a list of pairs, representing a successful parse each. The first component of the pair is the parsed value and the second component is the remaining input. The Haskell code for this definition is

```haskell
newtype Parser a = Parser (String -> [(a,String)])

parse :: Parser a -> String -> [(a,String)]
parse (Parser p) = p

instance Monad Parser where
  return x = Parser (\s -> [(x,s)])
  p >>= q  = Parser (\s ->
               concat [parse (q x) s' | (x,s') <- parse p s ])
```

where the monadic structure is defined by `bind` and `return`. Given a value, the `return` function creates a monad that consumes no input and simply returns the given value. The »= function acts as a sequencing operator for parsers. It takes two parsers and applies the second one over the remaining inputs of the first one, using the parsed values on the first parsing as arguments.

An example of primitive **parser** is the `item` parser, which consumes a character from a non-empty string. It is written in Haskell code as

```haskell
item :: Parser Char
item = Parser (\s -> case s of
                       "" -> []
                       (c:s') -> [(c,s')])
```

and an example of **parser combinator** is the `many` function, which allows one or more applications of the parser given as an argument

```haskell
many :: Paser a -> Parser [a]
many p = do
  a  <- p
  as <- many p
  return (a:as)
```

in this example `many item` would be a parser consuming all characters from the input string.

## 4.2 PARSEC

**Parsec** is a monadic parser combinator Haskell library described in [Lei01]. We have chosen to use it due to its simplicity and extensive documentation. As we expect to use it to parse user live input, which will tend to be short, performance is not a critical concern. A high-performace library supporting incremental parsing, such as **Attoparsec** [O'S16], would be suitable otherwise.

# 5 USAGE

## 5.1 JUPYTER KERNEL

The **Jupyter Project** [Tea] is an open source project providing support for interactive scientific computing. Specifically, the Jupyter Notebook provides a web application for creating interactive documents with live code and visualizations.

We have developed a Mikrokosmos kernel for the Jupyter Notebook, allowing the user to write and execute arbitrary Mikrokosmos code on this web application.

## 5.2 CODEMIRROR LEXER

Part IV

# TYPE THEORY

# 6

# INTUITIONISTIC LOGIC

## 6.1 CONSTRUCTIVE MATHEMATICS

## 6.2 THE DOUBLE NEGATION OF LEM IS PROVABLE

In intuitionistic logic, the double negation of the LEM holds for every proposition, that is,

$$\forall A \colon \neg\neg(A \vee \neg A)$$

- Proof Suppose $\neg(A \vee \neg A)$. We firstly are going to prove that, under this specific assumption, $\neg A$ holds. If $A$ were true, $A \vee \neg A$ would be true and we would arrive to a contradition, so $\neg A$. But then, if we have $\neg A$ we also have $A \vee \neg A$ and we arrive to a contradiction with the assumption. We should conclude that $\neg\neg(A \vee \neg A)$.
- Machine proof

```
id : {A : Set} → A → A
id a = a
```

Part V

# CONCLUSIONS

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