Termination checking in the $\lambda\Pi$-calculus modulo theory. From a practical and a theoretical viewpoint

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Termination checking in the λΠ-calculus modulo theory
From a practical and a theoretical viewpoint

Guillaume Genestier
From April 24th to August 4th 2017
Abstract

*Dedukti* is a type-checker for the λΠ-calculus modulo theory, which has the particularity to allow the user to declare rewrite rules, especially in order to encode the logic he/she wants to use. Thanks to the Curry-Howard-De Bruijn correspondence, the type checking done by *Dedukti* can be used to check proofs.

Unfortunately, to decide the type-checking, *Dedukti* needs the system of rewrite rules declared by the user, together with β-reduction, to be confluent and terminating. An external tool already exists to check the confluence, but no automatic tool exists for the termination, which must be checked independently by hand, before using *Dedukti*.

The subject of this report (and more widely of the internship) is to develop such a tool.

A first part of it is dedicated to an algorithm of termination checking: the *Size-Change Principle*. After a brief introduction to *Dedukti*, the algorithm is explained and the results obtained by an implementation of it for *Dedukti* are presented.

The second and longest part of this report is much more theoretic, since it presents reducibility candidates for the λΠ-calculus modulo theory. Those reducibility candidates can then be used together with the *Size-Change Principle* detailed in the first part, to check the termination of a rewrite rule system, as explained in [Wah07].
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Chapter 1

Implementing the Size-Change Principle for Deduki

1.1 What is Deduki

Deduki is a multipurpose type-checker based on the λΠ-calculus modulo theory [CD07].

The project Deduki was initiated by Gilles Dowek and is developed by the Inria-team Deducteam. A first version was developed by Matthieu Boespflug [Boe11] and Quentin Carbonneaux [BCH12] in Haskell and then in C. The current version of Deduki was developed mainly by Ronan Saillard [Sai15] and is programmed in OCaml.

The λΠ-calculus modulo theory is a calculus including only dependent types and rewrite rules. One can find the inference rules of this calculus, in this report section 2.2 or in the paper describing Deduki [ABC+16]. The rewrite rules allow the user to encode many logics in Deduki, making it multipurpose and peculiarly well-suited for interoperability between proof systems. This is the reason why a lot of programs have been developed to translate proofs from other proof assistants in Deduki and the way back. The graph below illustrate the interaction which exist between Deduki and a range of proof assistants or automatic provers.
Among the rules of λΠ-calculus modulo theory, there is one conversion rule:

$$
\frac{\Gamma \vdash t : A \quad \Gamma \vdash A : s \quad \Gamma \vdash B : s}{\Gamma \vdash t : B} \quad A \rightsquigarrow^* B
$$

Without going into the detail of the meaning of the rule, which can be found in section 2.2, this rule means that if we can rewrite the type $A$ into the type $B$ then every term of one type is of the other type too.

To decide type-checking, Dedukti must check if two types are joinable by rewriting. To do so, in Dedukti, the two types are reduced to their normal forms and then the equality of normal forms is checked syntactically. But this requires that the normal forms exist and will be reached, situation which is ensured if the rewrite rules declared by the user, together with β-reduction, are confluent and terminating.

In the current version of Dedukti, a tool exists to check confluence, but checking the termination stays at the charge of the user and out of the scope of Dedukti. Meaning that, when a proof is checked by Dedukti, the checking is under the condition the rules declared by the user are terminating.

There exists ad-hoc proofs that some rewrite rule systems are terminating, for instance [Dow17], but no tools actually check it automatically.

The first part of this report presents the implementation of a simple but quite efficient termination checking algorithm for Dedukti.
1.2 The Size-Change Principle

The Size-Change Principle is a criterion introduced by Lee, Jones and Ben-Amram in [LJBA01] for a first-order functional language.

The principle is to analyse how the arguments evolve between each call of a function. To perform this analysis, we must analyse the arguments of every application on the right-hand side of a rewrite rule.

1.2.1 An example to begin with

Let’s start the explanation on an example to get intuition and formalize it right after. For instance the Peano-like definition of the multiplication:

\[ \text{[n,n]} \text{ mult (S m) n} \rightarrow \text{plus n (mult m n)}. \]

There are two function applications on the right-hand side of this example:

- \( \text{plus} \) applied to \( n \) which is exactly the second argument of the left-hand side and \( \text{mult m n} \) which is bigger than both arguments of the left-hand side.
- \( \text{mult} \) applied to \( m \) which is a strict subterm of \( (S m) \) and \( n \) which is exactly the second argument of the left-hand side.

We will sum up all of this information on two matrices:

\[
\begin{array}{c|cc}
\text{plus} & \text{mult} & n \\
\hline
\text{mult} & (\infty & \infty) \\
\text{n} & = & \infty \\
\end{array}
\quad
\begin{array}{c|cc}
\text{mult} & m & n \\
\hline
\text{mult} & (S m) & < \\
\text{n} & \infty & = \\
\end{array}
\]

Where

- \( a < \) symbol means that the argument of the called function corresponding to the column is a strict subterm of the argument of the caller function corresponding to the row;
- \( a = \) symbol means that the argument of the called function and the argument of the caller function are the same;
- \( a \infty \) symbol means that the arguments have no direct link.

In the original article [LJBA01], the variation of the size of the arguments where encoded in bipartite graphs. The idea of summing up it in matrices, which are easier to manipulate with a programming language like OCaml comes from Rodolphe Lepigre and Christophe Raffalli in [LR16].

From the calls that are encoded in the rewrite rules, we can construct a call-graph. The arrows are annotated with the matrices that sum up the size-changes.

Here is the call-graph for our example:
In this case we know that the function \( \text{mult} \) is terminating, because, there is only one loop in the call-graph, which is indexed by \((< \infty)\) and there is a \(<\) on the diagonal of the matrix, meaning that between two recursive calls of the function, at least one argument (here the first one) has strictly decreased and the subterm order is well-founded.

**Definition 1.2.1 (Well-founded relation).** *We remind the reader that a relation \( R \) is well-founded on a class \( X \) if for all sets \( S \) in \( X \), if \( S \) is non-empty, \( S \) has a minimal element for \( R \). It can be rewrite :*

\[
\forall S \subseteq X, (S \neq \emptyset \Rightarrow \exists m \in S, \forall s \in S, (s, m) \notin R)
\]

*An other characterization, often more intuitive, is that a relation \( R \) is well-founded if there is no countable infinite descending chain.*

### 1.2.2 The tropical semiring and the transitive closure of call-graphs

Of course, the example we have chosen to present the *Size-Change Principle* is peculiarly easy, but in fact the criterion is quite efficient, especially to deal with mutual recursion and permuted arguments.

In those cases, the call-graph does not contain enough arrows to say if the rewrite rules system is terminating or not. Then, we have to compute the transitive closure of the call-graph. To do so, we have to multiply matrices containing \(<, =\) and \(\infty\), meaning that we must define a multiplication and an addition on such symbols.

In fact, those operations are the one of the *min-plus semiring* (often called the *tropical semiring*), with the values:

<table>
<thead>
<tr>
<th>symbol</th>
<th>value</th>
</tr>
</thead>
<tbody>
<tr>
<td>(&lt;)</td>
<td>(-1)</td>
</tr>
<tr>
<td>(=)</td>
<td>(0)</td>
</tr>
<tr>
<td>(\infty)</td>
<td>(\infty)</td>
</tr>
</tbody>
</table>

We then have the tables:

<table>
<thead>
<tr>
<th>+</th>
<th>&lt;</th>
<th>=</th>
<th>(\infty)</th>
</tr>
</thead>
<tbody>
<tr>
<td>&lt;</td>
<td>&lt;</td>
<td>&lt;</td>
<td>(\infty)</td>
</tr>
<tr>
<td>=</td>
<td>&lt;</td>
<td>=</td>
<td>=</td>
</tr>
<tr>
<td>(\infty)</td>
<td>&lt;</td>
<td>=</td>
<td>(\infty)</td>
</tr>
</tbody>
</table>

<table>
<thead>
<tr>
<th>(\times)</th>
<th>&lt;</th>
<th>=</th>
<th>(\infty)</th>
</tr>
</thead>
<tbody>
<tr>
<td>&lt;</td>
<td>&lt;</td>
<td>&lt;</td>
<td>(\infty)</td>
</tr>
<tr>
<td>=</td>
<td>&lt;</td>
<td>=</td>
<td>(\infty)</td>
</tr>
<tr>
<td>(\infty)</td>
<td>(\infty)</td>
<td>(\infty)</td>
<td>(\infty)</td>
</tr>
</tbody>
</table>
Those definitions of the operations are quite intuitive. When a function \( f_1 \) calls a function \( f_2 \) which calls a function \( f_3 \), we can link the arguments of \( f_3 \) to the ones of \( f_1 \) going through the ones of \( f_2 \).

Let’s choose an argument of \( f_1 \) and one of \( f_3 \) that we would like to link. To do so, we can use any of the argument of \( f_2 \) as a middle step, so it is possible to go through the best one, that justifies that the addition operation is \( \min \).

Once the best argument of \( f_2 \) is chosen, there is two links to multiply, one between \( f_1 \) and \( f_2 \) and the other on between \( f_2 \) and \( f_3 \).

• If a link is labelled \( \infty \) it means that we have no information about one part of the transformation, and we cannot recover it no matter what is the label of the other part.

• On the other hand, if a link is labelled \( < \) and the other one \( < \) or \( = \), we have a large and a strict decrease, so globally, we have a strict decrease.

• If both are labelled \( = \), the transitivity of the equality ensures us that the arguments of \( f_1 \) and \( f_3 \) are the same.

1.2.3 A definition of the Size-Change Principle

We now have all the ingredients, to give a definition of the Size-Change Principle.

Definition 1.2.2 (Fully applied terms). We consider a language which contains:

• variables: \( x, y \);

• constants: \( c, d \);

• functions: \( f, g \), each having an associated arity.

The set of fully applied terms is:

\[ t ::= x | c | f t_1 \ldots t_n, \text{ where } n \text{ is the arity of } f \]

Definition 1.2.3 (Size-Change Principle). We assume given a well-founded order on fully applied terms, which is used to construct the call-graph. The subterm order has this property and is the one chosen for the implementation.

A set of rewrite rules is Size-Change Terminating if:

• All rules are of the form

\[ f t_1 \ldots t_n \Rightarrow \tau \]

where \( n \) is the arity of \( f \) and \( t_1, \ldots, t_n, \tau \) are fully applied terms.

• The transitive closure of the call graph is such that all arrows linking a node with itself are labelled with a matrix having a \( < \) on the diagonal.
1.3 Implementation issues

1.3.1 Adapting an already existing implementation

When they developed PML, a program which is absolutely independent of Dedukti, Rodolphe Lepigre and Christophe Raffalli implemented the algorithm of the Size-Change Principle and they accepted to share their implementation with us.

This implementation provided us with the computation of the transitive closure of a call graph and the checking that every looping arrows are labelled with a matrix having a $<$ on the diagonal.

To do so efficiently, Rodolphe Lepigre and Christophe Raffalli introduced a notion of matrix subsuming another one.

**Definition 1.3.1** (Subsuming). A matrix $A = \begin{pmatrix} a_{1,1} & \ldots & a_{1,n} \\ \vdots & \ddots & \vdots \\ a_{m,1} & \ldots & a_{m,n} \end{pmatrix}$ subsumes the matrix $B = \begin{pmatrix} b_{1,1} & \ldots & b_{1,n} \\ \vdots & \ddots & \vdots \\ b_{m,1} & \ldots & b_{m,n} \end{pmatrix}$ if the two matrices have the same dimension and for every $i$ and $j$, $a_{i,j}$ is smaller than $b_{i,j}$. The subsuming relation is the pointwise extension of $<$.

**Property 1.3.2.** If a matrix $A$ subsumes a matrix $B$, then for all matrix $X$, when the dimensions are compatible, then $AX$ subsumes $BX$ and $XA$ subsumes $XB$.

**Proof.** It is direct, because the sum and the product are increasing functions in the tropical semiring. □

If two arrows link the same nodes and one is labelled with a matrix which subsumes the label of the other, then it is not useful to use the smallest matrix to compute the transitive closure. Indeed, when $A$ subsumes $B$, every $<$ symbol on the diagonal of $A$ is also on the diagonal of $B$ and, since the subsuming relation is stable by multiplication, if the Size-Change Principle is verified removing $B$ from the call-graph, it is also verified with both $A$ and $B$, since the new labels on the looping arrows contain at least as much $<$ on the diagonal.

1.3.2 The implementation and its results

Since the program developed for PML fitted perfectly our purposes for the call-graph computation, the aim was to develop a program interfacing Dedukti with it. Especially, it was necessary to identify the “calls” and to associate a matrix to them.

This goal led to the writing of 200 lines of OCaml which were added to the 300 from Rodolphe Lepigre and Christophe Raffalli’s program, making a complete working first prototype of termination checker for Dedukti.
Unfortunately, the Size-Change Principle, in its original version, was designed for a much poorer language than the λΠ-calculus modulo theory, including only functions applied to constructor patterns. Hence, most of the Dedukti files are out of the scope of this criterion, mostly because of the application of an argument.

For instance, Dedukti allows the user to declare:

```plaintext
List : Type.
Nil : List.
Cons : Nat -> List -> List.
def map : (Nat -> Nat) -> List -> List.
[ ] map _ Nil --> Nil.
[f, x, l] map f (Cons x l) --> Cons (f x) (map f l).
```

and

```plaintext
Lamt : Type.
Lam : (Lamt -> Lamt) -> Lamt.
[f, x] App (Lam f) x --> f x.
```

In the first example, the function `map` is terminating, whereas the second example is an encoding of the untyped λ-calculus, which is well-known for being non-terminating.

But those two examples are quite close in the point of view of the Size-Change Principle: we encounter a call to `f x` where `f` is a variable. A variable does not correspond with any nodes in the building call-graph. Then no called function can be identified and the program must raise an error since we have seen this situation can both occur in terminating and in non-terminating situations.

The implementation is working, but the criterion, in the form which has been implemented is far from being sufficient, since it accepts only some basics functions (like addition, multiplication, exponentiation, Ackerman function, list concatenation or list sorting...) and the encoding of the first-order logic restricted to implication, universal quantification and negation, sometimes called intuitionistic minimal logic.

This main problem led us to a more theoretic part of the internship, in which we had to enrich the criterion in order to make it more useful for our purpose of developing a termination checker for Dedukti. A few modifications of the criterion exists in the literature. We can for instance mention Pierre Hyvernat’s idea [Hyv13] to enrich the tropical semiring in order to cover the case where the size of an argument increases a little before decreasing more. Or Neil Jones and Nina Bohr [JB08] who modified the criterion to check termination of untyped λ-calculus terms.

Those enrichment were quite interesting, but it did not solve our main problem with the variable application, so we choose to investigate the work of David Wahlstedt [Wah07], who uses the Size-Change Principle in an original way and manage to study the termination of a system which is quite close of the λΠ-calculus modulo theory.
Chapter 2

Adaptation of the reducibility proof of Wahlstedt to the $\lambda\Pi$-calculus modulo theory

In his thesis [Wah07], David Wahlstedt proposes a calculus with Martin-Löf types and rewriting. He gives a criterion for the weak reducibility of a system of rewrite rules, coupled to the $\beta$-reduction. This work interested us, because it uses the Size-Change Principle in a rich system and it is very modular, allowing us to change only one part of the thesis to get the benefit of the entire one.

The structure of Wahlstedt’s thesis is the following:

- Define a reducibility predicate for weak normalisation.
- Show that if every function in the signature is reducible then every typable term is reducible.
- Show that if the call relation is well-founded and every type in the signature is reducible, then every function in the signature is reducible.
- The size-change principle is used to show that the call relation is well-founded.
- The reducibility of every type occurring in the signature is decidable.

Our aim in the second part of the internship was to adapt the tree first points to the $\lambda\Pi$-calculus modulo theory. Doing it, we can re-use the demonstration of the fourth point given in the original thesis and use the Size-Change Principle in a more efficient way.

2.1 Syntax

To begin, we will give the syntax of the $\lambda\Pi$-calculus modulo theory.
2.1.1 Terms and contexts

We use $x$ and $y$ to denote variables.

**Definition 2.1.1** (Terms).

\[
t, \tau, u, v, l, r ::= x | \lambda(x : u).t | tu | c | f
\]

We denote by $\Lambda$ the set of all terms.

**Definition 2.1.2** (Types).

\[
T, U ::= \lambda(x : U).T | \Pi(x : U).T | U v | d | F
\]

**Definition 2.1.3** (Kinds).

\[
K ::= Type | \Pi(x : U).K
\]

*Remark.* For now, $d, c$ and $f$ are only symbols. Their meaning will appear later in this section.

**Definition 2.1.4** (Contexts).

\[
\Gamma, \Delta ::= [ ] | \Gamma, x : t
\]

*Notations.* For readability, we write only the first $\Pi$ when they are chained.

\[
\Pi(x_1 : T_1)(x_2 : T_2)(x_3 : T_3)U \text{ means } \Pi(x_1 : T_1)(\Pi(x_2 : T_2)(\Pi(x_3 : T_3)U))
\]

We draw a line over the left part of a $\Pi$ or a $\lambda$, to mean that the declaration is iterated. The variable on which the iteration is done and the bounds are inferred by the reader from the context.

\[
\Pi\overline{x_i : T_i}U \text{ means } \Pi(x_1 : T_1) \ldots (x_n : T_n)U
\]

**Definition 2.1.5** (Closed context). A context $\Gamma = (x_1 : t_1, \ldots, x_i : t_i, \ldots)$ is closed if for all $i$, all free variables appearing in $t_i$ are in $\{x_1, \ldots, x_{i-1}\}$.

2.1.2 Substitutions

To express transformation that occur on terms during rewriting, it is necessary to define an operation of substitution, expressing the fact that every free occurrences of a particular variable takes the same value.

**Definition 2.1.6** (Substitution). More generally, a substitution $\delta$ a function which links every variable to a term. We define a function $\tilde{\delta}$ from terms to terms by applying simultaneously $\delta$ on every free variable. If one bound variable has the same name as a free variable occurring in the image of $\delta$, then the bound variable must be renamed, in order to avoid variable capture.
Remark. It is important to note that the linked variables are not affected by substitution. That prevents us from "capturing" variables that should not be.

Notations (Substitution). The notation to express it must mention the term in which the substitution occurs, the variable which is replaced and the term which replaces this variable, so it is quite natural to choose:

- \( \tau [t/x] \) to denote the term \( \tau \) in which \( x \) is replaced by \( t \).
- \( \delta, \sigma \) and \( \vartheta \) are used to denote a general substitution.
- We will denote by \( t\delta \) the substitution \( \tilde{\delta}(t) \) and by \( \delta(x) \) the image of the variable \( x \) by the function \( \delta \).

2.1.3 Signature

Definition 2.1.7 (Signature). A signature is a 6-tuple \( (\mathbb{D}, D, C, F, F) \).

- \( \mathbb{D} \) is a set of set constructors,
- \( D \) is a function which associates to each \( d \in \mathbb{D} \) a closed term of the form \( D(d) = \Pi(x_1 : T_1) \ldots (x_k : T_k) \) Type. Under this specification, \( k \) is called the arity of \( d \).
- \( C \) is a set of element constructors,
- \( C \) is a function which associates to each \( c \in C \) a closed term of the form \( C(c) = \Pi(x_1 : U_1) \ldots (x_m : U_m)(d\tau_1 \ldots \tau_k) \) where for all \( i \), \( \text{FV}(\tau_i) \subseteq \{x_1, \ldots, x_m\} \) and \( k \) is the arity of \( d \). Under this specification, \( m \) is called the arity of \( c \).
- \( F \) is a set of defined functions,
- \( F \) is a function which associates to each \( f \in F \) a closed term of the form \( F(f) = \Pi(x_1 : T_1) \ldots (x_n : T_n)U \) where \( U \) is not a \( \Pi \). Under this specification, \( n \) is called the arity of \( f \).

Notations. From now, \( d, c \) and \( f \) will always represent set constructor, element constructor and defined function respectively.

We denote \( \text{ar}(h) \) the arity of the symbol \( h \).

Remark. The arity of any symbol defined in the signature can be 0.

Definition 2.1.8 (Elementary types). We name elementary types a fully applied set constructor.

Definition 2.1.9 (Precedence on set constructors). We define the relation on \( \mathbb{D} \): \( \preceq_D \) as the reflexive-transitive cloure of \( d' \preceq_D d \) if:

- \( D(d) = \Pi(x_i : T_i) \) Type and \( d' \) appears in a \( T_i \)
- or if there is a \( c \) such that \( C(c) = \Pi(x_i : U_i)(d\tilde{s}) \) and \( d' \) appears in a \( U_i \).
Definition 2.1.10 ($\mathcal{D}$-equivalence). If $d \preceq_{\mathcal{D}} d'$ and $d' \preceq_{\mathcal{D}} d$, $d$ and $d'$ are said $\mathcal{D}$-equivalent, what is denoted $d \approx_{\mathcal{D}} d'$.

Remark. $\preceq_{\mathcal{D}}$ is by definition a pre-order, so $\approx_{\mathcal{D}}$ is an equivalence relation.

Since we only consider finite signature, $\preceq_{\mathcal{D}}$ is well-founded.

Definition 2.1.11 (Strictly positive set constructor). A set constructor $d$ is said strictly positive if:

- $\mathcal{D}(d) = \Pi(x_i : T_i)$. Type and no symbol $\mathcal{D}$-equivalent to $d$ occurs in any $T_i$.
- For all $c$ such that $\mathcal{C}(c) = \Pi(x_i : U_i). (d \bar{s})$, if the $U_i$ are of the form $\Pi(y_j : V_j). V$ then no symbol $\mathcal{D}$-equivalent to $d$ appears in a $V_j$.

We restrict our study to the signatures where every set constructor is strictly positive.

Definition 2.1.12 (Constructor pattern). We define this syntactical sub-category of terms as:

$$ p ::= x | cp_1 \ldots p_n $$

Definition 2.1.13 ($\beta$-normal term). We define this syntactical sub-category of terms as:

$$ s ::= xs_1 \ldots s_n | hs_1 \ldots s_{\text{ar}(h)} | \lambda(x : T). s $$

where $h$ is one symbol in the signature, set constructor, element constructor or defined function.

Now, we have all the necessary definitions to enrich the signature with rewrite rules.

Definition 2.1.14 (Rewrite rules in the signature). The signature is enriched with a set $\mathcal{R}$ of rewrite rules. Each rewrite rule is of the form $fp_1 \ldots p_k \rightarrow s$ where:

- the $p_i$ are constructor patterns,
- $k \leq \text{ar}(f)$,
- $p_k$ is not a variable,
- $s$ is $\beta$-normal,
- $s$ starts with $(\text{ar}(f) - k) \lambda$-abstractions,
- the rule is left-linear, meaning that a free variable can't appear twice in $fp_1 \ldots p_k$. 

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Furthermore, we demand \( R \) to be non-overlapping, meaning that for any two distinct rules \( l_1 \rightarrow r_1 \) and \( l_2 \rightarrow r_2 \) in \( R \), there are no substitutions \( \delta_1 \) and \( \delta_2 \) such that \( l_1 \delta_1 = l_2 \delta_2 \).

Remark. We do not have defined typable terms yet, but once it will be done, we will restrict our study to the rewrite rules that preserve typing.

2.2 Rules of inference

Now that we have presented the syntax of terms and contexts, we can introduce the rules of inference of the \( \lambda \Pi \)-calculus modulo theory.

Context formation

\[
\begin{array}{c}
\text{Well-formed} \\
\hline
\Gamma, x : A \quad \Gamma \vdash A : \text{Kind} \\
\hline
\Gamma, x : A \quad \Gamma \vdash A : \text{Type} \\
\hline
\end{array}
\]

Axioms

\[
\begin{array}{c}
\Gamma \vdash \text{Type} : \text{Kind} \\
\hline
\Gamma, x : A \quad \Gamma \vdash x : A \\
\hline
\end{array}
\]

Product

\[
\begin{array}{c}
\Gamma \vdash A : \text{Type} \\
\hline
\Gamma \vdash \Pi(x : A) B : \text{Kind} \\
\hline
\Gamma \vdash A, x : A \vdash B : \text{Type} \\
\hline
\end{array}
\]

Lambda-abstraction

\[
\begin{array}{c}
\Gamma \vdash A : \text{Type} \\
\hline
\Gamma \vdash \lambda(x : A). t : \Pi(x : A) B \\
\hline
\Gamma \vdash A, x : A \vdash B : \text{Type} \\
\hline
\Gamma \vdash A : \text{Type} \\
\hline
\Gamma \vdash \lambda(x : A). t : \Pi(x : A) B \\
\hline
\Gamma \vdash A, x : A \vdash B : \text{Type} \\
\hline
\end{array}
\]

Application

\[
\Gamma \vdash t : \Pi(x : A) B \\
\hline
\Gamma \vdash u : A \\
\hline
\Gamma \vdash t u : B [u/x]
\]

\(^1\)We have chosen here not to enforce variables to be distinct in each rules, since in practice the files considered in Dedukti do not follow this restriction. It would have been possible to consider only one substitution and to rename variables, which is equivalent to the presentation with two substitutions.
Signature symbols

\[ \frac{\Gamma, x_1 : T_1, \ldots, x_i : T_i \vdash T_{i+1} : \text{Type}}{\Gamma \vdash d : D(d)} \quad D(d) = \Pi(x_i : T_i) \text{ Type} \]

\[ \frac{\Gamma, x_1 : U_1, \ldots, x_i : U_i \vdash U_{i+1} : \text{Type}}{\Gamma \vdash c : C(c)} \quad C(c) = \Pi(x_i : U_i)(d \tau_1 \ldots \tau_k) \quad D(d) = \Pi(x_i : T_i) \text{ Type} \]

\[ \frac{\Gamma, x_1 : T_1, \ldots, x_i : T_i \vdash U_i \vdash T_{i+1} : \text{Type}}{\Gamma \vdash f : F(f)} \quad F(f) = \Pi(x_i : T_i)U \]

Conversion

\[ \frac{\Gamma \vdash t : A \quad \Gamma \vdash A : \text{Type} \quad \Gamma \vdash B : \text{Type}}{\Gamma \vdash t : B} \quad A \rightsquigarrow_* B \]

\[ \frac{\Gamma \vdash t : A \quad \Gamma \vdash A : \text{Kind} \quad \Gamma \vdash B : \text{Kind}}{\Gamma \vdash t : B} \quad A \rightsquigarrow_* B \]

2.2.1 Rewriting

**Definition 2.2.1** (Well-typed rules). *Now that we have given the typing rules of the \( \lambda \Pi \)-calculus, we can restrict our study to signatures which are well-typed, meaning that:

- all the rewrite rules in \( R \) preserve typing. For all \( \Gamma, t, T, u \), if \( \Gamma \vdash t : T \) and \( t \rightarrow u \in R \), then \( \Gamma \vdash u : T \).
- all the left-hand sides of rules are typable. For all \( f \bar{p} \rightarrow u \in R \) with \( F(f) = \Pi(x_i : T_i)U \) there is a \( \Delta_0 \) such that for all \( i \leq m \),

\[ \Delta_0 \vdash p_i : T_i[p_1/x_1 \ldots p_{i-1}/x_{i-1}] \]

When we add to the rewrite rules declared in \( R \) the \( \beta \)-reduction \( (\lambda x. t) u \rightarrow u [t/x] \), we can define the rewriting relation \( \rightsquigarrow \).

**Definition 2.2.2** (Rewriting relation). *We define inductively the relation, by :*

- if \( t \beta \)-reduces to \( u \) then \( t \rightsquigarrow u \),
- if there is \( l \rightarrow r \in R \) and \( \delta \) such that \( l \delta = t \) and \( r \delta = u \) then \( t \rightsquigarrow u \),
- if \( t \rightsquigarrow u \) then for all \( v \), \( t v \rightsquigarrow u v \), \( v t \rightsquigarrow v u \),
- if \( t \rightsquigarrow u \) then \( \lambda x. t \rightsquigarrow \lambda x. u \) and \( \Pi x. t \rightsquigarrow \Pi x. u \).
Notations. We denote by $\sim^*$ the reflexive-transitive closure of $\sim$ and by $\leftrightarrow^*$ the reflexive-symmetric-transitive closure of $\sim$.

**Definition 2.2.3** (Confluence). A system of rewrite rules is confluent if for any terms $t$, $t_1$ and $t_2$, if $t \sim^* t_1$ and $t \sim^* t_2$ then there is a $u$ such that $t_1 \sim^* u$ and $t_2 \sim^* u$.

**Proposition 2.2.4.** With all the constraints on rewrite rules, which are given by definition 2.1.14 and the preservation of typing, we can ensure that $R$ is confluent.

We won’t prove this proposition, which is a consequence of orthogonality [OR94], because our purpose is to study termination and not confluence. The interested reader can find the demonstration, which is quite long and does not introduce techniques that will be useful anywhere else in this report, in [Wah07, section 2.2.3 page 31]

### 2.3 Reducibility predicate

#### 2.3.1 Weak normalisation

**Definition 2.3.1** (Normalisation predicates).

\[
\begin{align*}
\text{NF}(u) &\equiv \neg (\exists v. u \sim v) \\
\text{SN}(u) &\equiv \neg (\exists (v_i)_{i \in \mathbb{N}}. v_0 = u \land \forall i. v_i \sim v_{i+1}) \\
\text{WN}(u) &\equiv \exists v. u \sim^* v \land \text{NF}(v) \\
\end{align*}
\]

**Definition 2.3.2** (Strongly neutral terms).

\[
b ::= x t_1 \ldots t_n \text{ where } \text{NF}(t_i)
\]

\[
| f t_1 \ldots t_n \text{ where } \text{NF}(f t_1 \ldots t_n) \text{ and } n \geq \text{ar}(f)
\]

We denote by $\mathcal{N}$ the set of strongly neutral terms.

**Lemma 2.3.3** (Weak normalisation of an application). If $x$ is a variable and $t$ a term, then WN($t x$) $\Rightarrow$ WN($t$)

**Proof.** Assume WN($t x$). Then there is a reduction sequence

\[
t x \sim t_1 \sim t_2 \sim \ldots \sim t_n
\]

where $t_n$ is normal. We proceed by induction on the length $n$ of the shortest reduction.

- If $n = 0$ then $t x$ is normal, then $t$ is normal too.
- Else $t x$ contains a redex. Applying a term to a variable cannot be of the form $l \delta$ where $l$ is the left-hand side of a rule and $\delta$ a substitution, so the first step $t x \sim t_1$ must be of the form :
\[(\lambda y.u)x \to u[x/y] \text{ and since } u[x/y] \text{ is normalizable, } u \text{ is normalizable too, hence } \text{WN}(t).\]

\[tx \to u x, \text{ because } t \to u. \text{ Then } u x \text{ reduce to } t_n \text{ so we have } \text{WN}(u x) \text{ and there is a reduction in } n - 1 \text{ steps so by the induction hypothesis, we have } \text{WN}(u). \text{ We can deduce } \text{WN}(t) \]

2.3.2 Reducible terms

**Definition 2.3.4 (Reducibility of a term).** We define mutually the predicates RED for Kind, the inhabitants of Kind and the inhabitants of those inhabitants.

\[\text{RED}_{(\text{Type})}(t) \text{ holds if one of the following conditions occurs:}\]

- \[t = \text{Type} \text{ and then } \text{RED}_{(\text{Type})}(U) \text{ holds if one of the following conditions occurs:}\]

\[- \exists d, u_1, \ldots, u_k. D(d) = \Pi(x_1 : T_1) \ldots (x_k : T_k) \text{ Type and then } \text{RED}(U)(v) \]

\[\forall i. \text{RED}_{(\text{Type})}(T_i) \land \text{RED}_{(U)}(u_i) \]

\[\text{holds if one of the following conditions occurs:}\]

\[- \exists c, v_1, \ldots, v_m. C(c) = \Pi(x_1 : U_1) \ldots (x_m : U_m)(d \tau_1 \ldots \tau_k) \]

\[\forall i. \text{RED}_{(\text{Type})}(U_i[v_1/x_1, \ldots, v_{i-1}/x_{i-1}]) \]

\[\forall i. \text{RED}_{(U)}(v_i) \]

\[\exists b \in \mathcal{N}. v \downarrow b \quad \exists b \in \mathcal{N}. U \downarrow b \quad \text{and then } \text{RED}(U)(v) \]

\[- \exists A, B. \begin{cases} t = \Pi(x : A) & B \\ \text{RED}_{(\text{Type})}(A) & \text{RED}_{(\text{Type})}(B[a/x]) \end{cases} \]

\[\forall a. \text{RED}_{(A)}(a) \Rightarrow \text{RED}_{(\text{Type})}(B[a/x]) \]

\[\text{and then } \text{RED}_{(\Pi(x:A)B)}(u) \]

\[- \exists b \in \mathcal{N}. U \downarrow b \text{ and then } \text{RED}_{(U)}(v) \text{ holds if } \exists b' \in \mathcal{N}. v \downarrow b' \]

- \[\exists A, B. \begin{cases} t = \Pi(x : A) & B \\ \text{RED}_{(\text{Type})}(A) & \text{RED}_{(\text{Type})}(B[a/x]) \end{cases} \]

\[\forall a. \text{RED}_{(A)}(a) \Rightarrow \text{RED}_{(\text{Type})}(B[a/x]) \]

\[\text{and then } \text{RED}_{(\Pi(x:A)B)}(u) \]

\[\exists A, B. \begin{cases} t = \Pi(x : A) & B \\ \text{RED}_{(\text{Type})}(A) & \text{RED}_{(\text{Type})}(B[a/x]) \end{cases} \]

\[\forall a. \text{RED}_{(A)}(a) \Rightarrow \text{RED}_{(\text{Type})}(B[a/x]) \]

\[\text{and then } \text{RED}_{(\Pi(x:A)B)}(u) \]

\[\text{Are the predicates RED well-defined ?}\]

We will ensure that the predicates RED are well-defined for elementary types by defining inflating sequences of sets and name RED their least fixpoints.
Definition 2.3.5 (Elementary type interpretations). Let $A = \{d_i\}_{i \in \{1, \ldots, n\}}$ be a $\approx_D$-equivalence class.

\[
F_A : \prod_{i=1}^n \mathcal{P}(A) \to \prod_{i=1}^n \mathcal{P}(A)
\]

\[
(X_i)_{i \in \{1, \ldots, n\}} \mapsto \left\{ u \mid \exists b. u \downarrow b \right\} \cup \left\{ u \mid \exists c, v_1, \ldots, v_m. C(c) = \Pi(x_j : U_j). (d_i \bar{s}) \forall j. v_j \in R_{U_j}(X) \right\}
\]

where $R_\alpha(X) = \begin{cases} \mathbb{[}d]\text{ if } \alpha = d \bar{s}, \text{ where } d \text{ is fully applied and } \forall u \in A. d \prec_D u & \\
X_i \text{ if } \alpha = d_i \bar{s}, \text{ where } d \text{ is fully applied} & \\
N \text{ if } \alpha \in \mathcal{N} & \\
\{ t \in \Lambda \mid \forall u \in R_{T_1}(X), t u \in R_{T_2[u/\bar{x}]}(X) \} \text{ if } \alpha = \Pi(x : T_1). T_2 & \end{cases}
$

where, for all $\approx_D$-equivalence class $A'$ where $A' \prec_D A$ and $d \in A'$, $\mathbb{[}d]\text{ is the least fixpoint of } F_{A'}$.  

Remark. We must insist here on the fact that $\mathbb{[}d]\text{ must be seen as the interpretation of any elementary type of the form } d t_1 \ldots t_n \text{ where } d \text{ is fully applied.}$

Existence of the least fix point. We will show that $F_A$ is increasing. Thus, from Knaster-Tarski theorem, the function has a least fixpoint. Let’s take $X \subseteq \prod \bar{Y} \subseteq \Lambda$ and $u \in F_A(X)$. Two cases occur:

- $u \downarrow b$, then $u \in F_A(\bar{Y})$
- $\exists c. u \downarrow c \bar{v}, C(c) = \Pi(x_j : U_j). (d \bar{s})$ and $\forall j. v_j \in R_{U_j}(X)$. Then two subcases appear:

  - $U_j = \Pi(\bar{x}_k : T_k). d' \bar{\tau}$ such that for all $d_i \in A$ we have $d' \prec_D d_i$. As every elementary type is strictly positive, all the symbols occurring in the $T_k$ are strictly smaller for the order $\prec_D$ than those in $A$, so for all $k$, $R_{T_k}(\bar{Y}) = R_{T_k}(X)$. For $\tau_k \in R_{T_k}(X)$ then $v_j \bar{\tau} \in \mathbb{[}d']$ so $v_j \in R_{U_j}(\bar{Y})$.

  - $U_j = \Pi(\bar{x}_k : T_k). d_i \bar{\tau}$. As every elementary type is strictly positive, all the symbols occurring in the $T_k$ are strictly smaller for the order $\prec_D$ than those in $A$, so $R_{T_1}(\bar{Y}) = R_{T_1}(X)$. For $\tau_k \in R_{T_k}(X)$ then $v_j \bar{\tau} \in X_i$, then $v_j \bar{\tau} \in Y_i$ because $X \subseteq \prod \bar{Y}$ so $v_j \in R_{U_j}(\bar{Y})$.

So for all $j$, we have $v_j \in R_{U_j}(\bar{Y})$, hence, $u = c \bar{v}$ is in $F_A(\bar{Y})$

So $F_A(\bar{X}) \subseteq F_A(\bar{Y})$, hence $F_A$ is increasing.

Hence, since $\prec_D$ is well founded, we proceed inductively on the $\approx_D$ equivalence classes to define the interpretation of every elementary types.

Definition 2.3.6 (Reducible terms). We can define RED$_T(u)$, where $T$ reduces to an elementary type by

$\exists d, u_1, \ldots, u_m. T \downarrow d u_1 \ldots u_m$ and $u \in \mathbb{[}d]$. 

We can define \( \text{RED}_T(u) \), where \( T \) reduces to a strongly normal type by \( u \in \mathcal{N} \).

Then we can define, by induction on the number of \( \Pi \) in the normal form, \( \text{RED}_T(u) \), where \( T \) reduces to \( \Pi(x : A)U \) by \( \{ t \in \Lambda | \forall u \in [A].t u \in [U [u/x]] \} \).

**Definition 2.3.7 (Reducible types).** Once reducible terms are well-defined, we can define reducible types by the increasing sequence:

- \( \text{Type}(0) = \{ U | \exists b. U \downarrow b \} \)
- For any ordinal \( \alpha \),
  \[
  \text{Type}(\alpha + 1) = \text{Type}(\alpha) \cup \left\{ U \mid \exists d, u_1, \ldots, u_m. D(d) = \Pi(x_1 : T_1) \ldots (x_k : T_k) \text{ Type} \right\} \]
  \[
  \forall i. T_i \in \text{Type}(\alpha) \wedge \text{RED}(T_i)(u_i) \]
- For any limit \( \lambda \),
  \[
  \text{Type}(\lambda) = \bigcup_{\alpha < \lambda} \text{Type}(\alpha) \]

Then \( \text{RED}_{\text{Type}}(T) \) if and only if \( T \) is in the least fixpoint of \( (\text{Type}(\alpha))_\alpha \).

**Definition 2.3.8 (Reducible kinds).** Similarly, we then can define reducible kinds by the increasing sequence:

- \( \text{Kind}(0) = \{ \text{Type} \} \)
- For any ordinal \( \alpha \),
  \[
  \text{Kind}(\alpha + 1) = \text{Kind}(\alpha) \cup \left\{ \Pi(x : A) B \mid \forall a. \text{RED}(A)(a) \Rightarrow B \left[ a/x \right] \in \text{Kind}(\alpha) \right\} \]
- For any limit \( \lambda \),
  \[
  \text{RED}(\lambda) = \bigcup_{\alpha < \lambda} \text{RED}(\alpha) \]

Then \( \text{RED}_{\text{Kind}}(K) \) if and only if \( K \) is in the least fixpoint of \( (\text{Kind}(\alpha))_\alpha \).
Properties of reducibility

**Proposition 2.3.9** (Conversion does not change reducibility). If $T \rightsquigarrow^* T'$ and $t \rightsquigarrow^* t'$ then RED$_{(T)}(t) \Leftrightarrow$ RED$_{(T')}(t')$

Proof. Since the reducibility predicate definition, does a matching on disjoint cases considering only the normal form of a term and since the rewrite rules considered are confluent, this property holds. □

**Proposition 2.3.10.** For all $T$, $t$, if RED$_{(\text{Kind})}(T)$ then

1. RED$_{(T)}(t) \Rightarrow$ WN$(t)$
2. $t \downarrow b \Rightarrow$ RED$_{(T)}(t)$

Proof. We prove both mutually by induction.

1. • If $T =$ Type it works by definition of the reducibility predicate.
   • If $T =$ Π$(x : U)V$. Suppose RED$_{(\text{Kind})}($Π$(x : U)V$), we have RED$_{(\text{Type})}(U)$ and $(\forall u, \text{RED}$_{(U)}(u) \Rightarrow \text{RED}$_{(\text{Kind})}(V[u/x]))$. Suppose we furthermore have RED$_{(\Pi(x:U)V)}(t)$, we can then deduce $(\forall u, \text{RED}$_{(U)}(u) \Rightarrow \text{RED}$_{(V[u/x]})(t u))$. Since RED$_{(\text{Type})}(U)$ and $x$ is strongly neutral, by induction case 2, we obtain RED$_{(U)}(x)$ so RED$_{(\text{Kind})}(V)$ and RED$_{(V)}(t x)$. We can deduce WN$(t x)$ by induction and then WN$(t)$ by lemma 2.3.3.

2. • If $T =$ Type it works by definition of the reducibility predicate.
   • If $T =$ Π$(x : U)V$. Suppose RED$_{(\text{Kind})}($Π$(x : U)V$), we have RED$_{(\text{Type})}(U)$ and $(\forall u, \text{RED}$_{(U)}(u) \Rightarrow \text{RED}$_{(\text{Kind})}(V[u/x]))$. We take a $u$ such that RED$_{(U)}(u)$. By induction case 1, there is a $u'$ such that $t \downarrow u'$ and by definition $b u'$ is strongly neutral. So by induction case 2, RED$_{(\Pi(x:U)V)}(t u)$ because $t u \downarrow b u'$ which is strongly neutral. So RED$_{(\Pi(x:U)V)}(t)$ by definition of the reducibility predicate. □

**Proposition 2.3.11.** For all $T$, $t$, if RED$_{(\text{Type})}(T)$ then

1. RED$_{(T)}(t) \Rightarrow$ WN$(t)$
2. $t \downarrow b \Rightarrow$ RED$_{(T)}(t)$

Proof. We prove both mutually by induction.

1. • If $T \downarrow b$ or $T \downarrow d_t_1 \cdots t_n$ it works by definition of the reducibility predicate.
   • If $T \downarrow \Pi(x : U)V$. Suppose RED$_{(\text{Type})}($Π$(x : U)V$), we have RED$_{(\text{Type})}(U)$ and $(\forall u, \text{RED}$_{(U)}(u) \Rightarrow \text{RED}$_{(\text{Type})}(V[u/x]))$. Suppose we furthermore have RED$_{(\Pi(x:U)V)}(t)$, we can then deduce $(\forall u, \text{RED}$_{(U)}(u) \Rightarrow \text{RED}$_{(V[u/x]})(t u))$. Since RED$_{(\text{Type})}(U)$ and $x$
is strongly neutral, by induction case 2, we obtain RED\(_{(U)}(x)\) so RED\(_{(Type)}(V)\) and RED\(_{(V)}(t.x)\). We can deduce WN(t.x) by induction and then WN(t) by lemma 2.3.3.

2. • If T \Downarrow b\ or T \Downarrow d t_1 \ldots t_n\ it works by definition of the reducibility predicate.

• If T \Downarrow \Pi(x : U)V. Suppose RED\(_{(Type)}(T)\)\, we have RED\(_{(Type)}(U)\) and (\(\forall u, RED\(_{(U)}(u) \Rightarrow RED\(_{(Type)}(V[u/x])\)\)). We take a u such that RED\(_{(U)}(u)\). By induction case 1, there is a u' such that u \Downarrow u' and by definition b u' is strongly neutral. So by induction case 2, RED\(_{(V[u/x]}(t u)\) because t u \Downarrow b u' which is strongly neutral. So RED\(_{(\Pi(x : U)V)}(t)\) by definition of the reducibility predicate. \(\Box\)

2.3.3 Reducibility in a context

**Definition 2.3.12** (Reducibility of a substitution). RED\(_{(x_1 : T_1, \ldots, x_n : T_n)}(\gamma)\) holds whenever \(\forall i, RED\(_{(Type)}(T_i, \gamma) \land RED\(_{(T, \gamma)}(\gamma(x_i))\)\).

**Proposition 2.3.13.** If \(\Gamma\) is closed, RED\(_{(\Gamma, x : T)}(\gamma)\), RED\(_{(T, \gamma)}(t)\), \(x \notin supp(\Gamma)\) and \(x \notin FV(T)\), then RED\(_{(\Gamma, x : T)}([\gamma, t/x])\)

**Proof.** Assume given \((y : U)\) in \((\Gamma, x : T)\). We have two cases:

- \(x = y\) Then \(T = U\) and \(y[y, t/x] = x[y, t/x] = t\) and \(T[y, t/x] = T\gamma\) because \(x\) is not free in \(T\). By assumption we have RED\(_{(T, \gamma)}(t)\) thus we have RED\(_{(T[y, t/x]}(y[y, t/x])\)

- \(x \neq y\) Then \(y[y, t/x] = y\gamma\). From RED\(_{(\Gamma, x : T)}(\gamma)\) we have RED\(_{(U, \gamma)}(\gamma(y))\) and we have \(U[y, t/x] = U\gamma\) because \(\Gamma\) is closed and \(x\) is not free in \(U\). So RED\(_{(U[y, t/x]}(y[y, t/x])\). \(\Box\)

2.3.4 Reducibility of well-typed terms

**Definition 2.3.14.** Let RED\(_{(\mathcal{F})}\) be the property:

For all \(f\), if \(\mathcal{F}(f)\) is of the form \(\Pi(x_i : T_i)\) Type then we have RED\(_{(Kind)}(\mathcal{F}(f))\) else we have RED\(_{(Type)}(\mathcal{F}(f))\)

**Lemma 2.3.15.** If RED\(_{(\mathcal{F})}\) holds, then, for all \(f\) such that \(f\) has no rewrite rule, we have RED\(_{(\mathcal{F}(f))}(f)\).

**Proof.** Let \(\mathcal{F}(f) = \Pi(x_i : T_i)U\). Take \(\tilde{t}_i\) such that for all \(i, RED\(_{(\gamma, t_1/x_1, \ldots, t_{i-1}/x_{i-1})}(\gamma)\)

By RED\(_{(\mathcal{F})}\), we have RED\(_{(\mathcal{S}ort)}(U[t_1/x_1])\) where Sort is Type or Kind depending on the case in which \(f\) is. By proposition 2.3.10 there are \(\tilde{u}_i\) such that \(\tilde{t}_i \Downarrow \tilde{u}_i\). And since \(f\) has no rewrite rule, \(f u_1 \ldots u_n\) is strongly neutral. Then, using the second case of proposition 2.3.10 we obtain RED\(_{(U[t_1/x_1]}(f u_1 \ldots u_n)\).
Finally, we can conclude \( \text{RED} \left( u \left[ t_1/x_1 \right] \right) (f t_1 \ldots t_n) \) using proposition 2.3.9.

\[ \square \]

**Theorem 2.3.16.** If \( \text{RED}(F) \) and \( \forall f. \text{RED}(F(f)) (f) \) then

\[ \Gamma \vdash t : T \Rightarrow [\forall \gamma. \text{RED}_T(\gamma) \Rightarrow \text{RED}_{T(\gamma)} (t \gamma)] \]

**Proof.** We will prove by induction on the typing derivation that :

\[ \forall \gamma. \text{RED}_T(\gamma) \Rightarrow \text{RED}_{T(\gamma)} (t \gamma) \]

Assume \( \Gamma \vdash t : T \) and given a \( \gamma \) such that \( \text{RED}_T(\gamma) \).

- If \( \Gamma \) well-formed

  \[ \frac{\Gamma \vdash \text{Type} : \text{Kind}}{\Gamma \vdash A : \text{Type}} \]

  By definition, we have \( \text{RED}_{(\text{Kind})} (\text{Type}) \).

- If \( \Gamma \) well-formed

  \[ \frac{\Gamma \vdash x : A}{x : A \in \Gamma} \]

  We have \( \text{RED}_T(\gamma) \), meaning that we have \( \text{RED}_{(A \gamma)} (\gamma(x)) \) for every \( (x : A) \in \Gamma \).

- \( \Gamma \vdash A : \text{Type} \quad \Gamma, x : A \vdash B : \text{Sort} \)

  \[ \frac{\varphi \in (\text{Sort})}{\Gamma \vdash \Pi(x : A) B : s} \]

  Sort is Kind or Type

  By induction hypothesis, we have \( \text{RED}_{(\text{Type})} (A \gamma) \). Let’s take a \( a \) such that \( \text{RED}_{(A \gamma)} (a) \). By proposition 2.3.13, we have \( \text{RED}_{(\Gamma, x : A)} (\gamma, a/x) \) and by induction \( \text{RED}_{(\text{Sort})} (B [\gamma, a/x]) \).

  It is exactly the definition of \( \text{RED}_{(\text{Sort})} ((\Pi(x : A) B) \gamma) \).

- \( \Gamma \vdash A : \text{Type} \quad \Gamma, x : A \vdash B : \text{Sort} \)

  \[ \frac{\varphi \in (\text{Sort})}{\Gamma \vdash \lambda(x : A).t : \Pi(x : A) B} \]

  Sort is Kind or Type

  By induction hypothesis, we have \( \text{RED}_{(\text{Type})} (A \gamma) \). Let’s take a \( a \) such that \( \text{RED}_{(A \gamma)} (a) \). By proposition 2.3.13, we have \( \text{RED}_{(\Gamma, x : A)} (\gamma, a/x) \) and by induction \( \text{RED}_{(\text{Sort})} (u [\gamma, a/x]) \) for every \( u \). Furthermore, \( (\lambda(x : A).t) a \sim t [a/x] \) and \( (((\lambda(x : A).t) a) \gamma) \sim t [\gamma, a/x] \). Then, by proposition 2.3.9, we can conclude that we have \( \text{RED}_{(\Pi(x : A) B) \gamma} (((\lambda(x : A).t) a) \gamma) \).

  It is exactly the definition of \( \text{RED}_{(\Pi(x : A) B) \gamma} (((\lambda(x : A).t) a) \gamma) \).

- \( \Gamma \vdash t : \Pi(x : A) B \quad \Gamma \vdash u : A \)

  \[ \frac{\varphi \in (\text{Sort})}{\Gamma \vdash t u : B [u/x]} \]

  By induction hypothesis, we have \( \text{RED}_{(\Pi(x : A) B) \gamma} (t \gamma) \) and \( \text{RED}_{(A \gamma)} (u \gamma) \).

  Then, by definition of the reducibility of a \( \Pi \) we have \( \text{RED}_{(B [u/x]) \gamma} ((t u) \gamma) \).

- \( \Gamma, x_1 : T_1, \ldots, x_i : T_i \vdash T_{i+1} : \text{Type} \)

  \[ \frac{\varphi \in (\text{Sort})}{\Gamma \vdash d : D(d)} \]

  \[ D(d) = \Pi(x_1 : T_1) \text{ Type} \]
Assume given \( t_i \) such that RED \( \left\{ t, \gamma; t_1/x_1, \ldots, t_{i-1}/x_{i-1} \right\} (t_i) \). By proposition 2.3.10, there exist \( u_i \) such that \( t_i \downarrow u_i \). By proposition 2.3.9, we have RED \( \left\{ t, \gamma; t_1/x_1, \ldots, t_{i-1}/x_{i-1} \right\} (u_i) \). Furthermore, by induction hypothesis, we have RED\((\text{Type})\) \( (T_i \gamma) \).

Hence, we obtain RED\((\text{Type})\) \( (d t_1 \ldots t_k) \), and this is exactly the definition of RED\((\Pi(x_1:T_1) \ldots (x_k:T_k) \text{ Type}) \gamma) \( (d) \)

\[ \Gamma, y_1 : U_1, \ldots, y_i : U_i \vdash U_{i+1} : \text{Type} \quad \Gamma, x_i : U_i \vdash T_j \quad \Gamma \vdash c : C(c) \]

\[ \frac{}{\Gamma \vdash f : F(f)} \]

By hypothesis, we have RED\((F(f)) \gamma (f) \). Since \( F(f) \) is closed, we have RED\((\text{Type}) \gamma (f \gamma) \).

\[ \Gamma, x_1 : T_1, \ldots, x_i : T_i \vdash T_{i+1} : \text{Type} \quad \Gamma \vdash f : F(f) \quad \Gamma \vdash A \quad \Gamma \vdash B \quad \Gamma \vdash t : : 2 \cdot 2 \cdot 2 \cdot 2 \]

\[ \frac{}{\Gamma \vdash t : A \wedge A : \text{Sort} \quad \Gamma \vdash B : \text{Sort} \quad \Gamma \vdash t : B \quad \text{Sort is Kind or Type}} \]

By induction, we have RED\((A \gamma) (t \gamma) \) and by proposition 2.3.9, we obtain that RED\((B \gamma) (t \gamma) \), because from \( A \sim B \), we have \( A \sim B \gamma \).

\[ \Box \]

2.4 Call-relation and reducibility of defined functions

2.4.1 Call relation

**Definition 2.4.1** (Formal call). We define \( (f, (p_1, \ldots, p_m)) \succ (g, (u_1, \ldots, u_n)) \) by:

- there is a \( k \) such that \( f p_1 \ldots p_k \rightarrow s \) is in \( \mathbb{R} \),
- \( \text{ar}(f) = m, \text{ar}(g) = n \),
• $gu_1 \ldots u_n$ is a subterm of $sp_{k+1} \ldots p_m$.

These types of calls are the one on which the Size-Change Principle algorithm, described section 1.2, will be performed. However, we will use another notion of call to state some interesting results about our reducibility predicates.

**Definition 2.4.2 (Instantiated call).** $(f, (t_1, \ldots, t_m)) \preceq (g, (u_1, \ldots, u_n))$ holds if there exist $p_1, \ldots, p_m, u_1, \ldots, u_n$ and a substitution $\gamma$ such that:

- $(f, (p_1, \ldots, p_m)) \succ (g, (u_1, \ldots, u_n))$,
- $\forall i \leq m. t_i \leadsto p_i \gamma$,
- $\forall i \leq m. \WN(t_i)$,
- $\forall j \leq n. v_j = u_j \gamma$.

### 2.4.2 Proof of reducibility for defined functions

**Theorem 2.4.3.** If $\succeq$ is well-founded and $\text{RED}(\mathcal{F})$ then

$$\forall f. \text{RED}(\mathcal{F}(f)) (f)$$

**Proof.** As the hypothesis suggests, we want to perform a well-founded induction on $\succeq$. But some work is required before doing this induction.

For all $f \in \mathcal{F}$, let’s write

$$\mathcal{F}(f) = \Pi(x_1 : T_{f,1}) \ldots (x_{\text{ar}(f)} : T_{f,\text{ar}(f)}) U_f$$

We will denote by $\Phi(f, \bar{t})$ the property:

$$\forall i. \text{RED} \left( T_{f,i} \left[ t_1/x_1, \ldots, t_{i-1}/x_{i-1} \right] \right) (t_i) \Rightarrow \text{RED} \left( U_{f,i} \left[ t_i/x_i \right] \right) (f \bar{t})$$

Here, we should note that $\forall f. \forall \bar{t}. \Phi(f, \bar{t})$ is the definition of $\forall f. \text{RED}(\mathcal{F}(f)) (f)$ which is our goal.

Assume given $f$ and $t_1, \ldots, t_{\text{ar}(f)}$ such that

$$\forall i. \text{RED}_{\text{Type}} \left( T_{f,i} \left[ t_1/x_1, \ldots, t_{i-1}/x_{i-1} \right] \right)$$

$$\forall i. \text{RED} \left( T_{f,i} \left[ t_1/x_1, \ldots, t_{i-1}/x_{i-1} \right] \right) (t_i)$$

$$\forall g. \forall \bar{u}. (f, \bar{t}) \succeq (g, \bar{u}) \Rightarrow \Phi(g, \bar{u})$$

From proposition 2.3.10, there exist $\bar{v}$ such that $\bar{t} \downarrow \bar{v}$. Then we have two cases:

- $f \bar{v}$ is normal, then since $f$ is fully applied, $f \bar{v}$ is strongly neutral and applying the other case of the proposition 2.3.10 we obtain $\text{RED} \left( U_{f,i} \left[ t_i/x_i \right] \right) (f \bar{t})$
Otherwise \( f \bar{v} \) matches a rewrite rule declared by the user. So there are a 
\( k \), a rewrite rule \( f p_1 \ldots p_k \rightarrow s \) and a substitution \( \vartheta \) such that \( \forall i \leq k. v_i = p_i \vartheta \). Then \( s = \lambda (x_{k+1} : T_{k+1}) \ldots (x_{\ar(f)} : T_{\ar(f)}). s_0 \) and

\[
f v_1 \ldots v_{\ar(f)} \rightsquigarrow^* s_0 \left[ \vartheta, v_{k+1}/x_{k+1}, \ldots, v_{\ar(f)}/x_{\ar(f)} \right]
\]

We will denote \( \gamma \) the substitution \( \left[ \vartheta, v_{k+1}/x_{k+1}, \ldots, v_{\ar(f)}/x_{\ar(f)} \right] \) and prove later that \( \text{RED} \left( U_f \left[ t_i/x_i \right] \right) (s_0 \gamma) \). Once we have it, by proposition 2.3.9, we obtain \( \text{RED} \left( U_f \left[ t_i/x_i \right] \right) (f \bar{t}) \)

In both cases, we can conclude \( \Phi(f, \bar{t}) \), then we have:

\[
\forall f. \forall \bar{t}. (\forall g. \forall \bar{u}. (f, \bar{t}) \leadsto (g, \bar{u}) \Rightarrow \Phi(g, \bar{u})) \Rightarrow \Phi(f, \bar{t})
\]

and then by the induction principle on the well-founded relation \( \leadsto \), we can conclude \( \forall f. \forall \bar{t}. \Phi(f, \bar{t}) \) which is the definition of

\[
\forall f. \text{RED}(\mathcal{F}(f)) (f)
\]

**Remaining goal 1.** We will now show that if \( f \bar{t} \) matches a rewrite rule declared by the user, we have

\[
\text{RED} \left( U_f \left[ t_i/x_i \right] \right) (s_0 \gamma).
\]

**Proof.** Since we have supposed that the left-hand side of all the rewrite rules are typable, there is a \( \Delta_0 \) such that:

\[
\forall i \leq m. \Delta_0 \vdash p_i : T_f, i \left[ p_1/x_1 \ldots p_{i-1}/x_{i-1} \right]
\]

\[
\forall m < i \leq \ar(f). \Delta_0 \vdash x_i : T_f, i \left[ p_1/x_1 \ldots p_m/x_m \right]
\]

Since all the rewrite rules preserve typing, we know that \( \Delta_0 \vdash s_0 : U_f \left[ \bar{p}/x_i \right] \).

We will show that for all subterms \( s \) of \( s_0 \), all contexts \( \Theta \) extending \( \Delta_0 \), all \( U \) and all substitutions \( \sigma \) which are equal to \( \gamma \) on \( \Delta_0 \), we have:

\[
(\Theta \vdash s : U \land \text{RED}_\Theta(\sigma) \land \text{RED}_{\text{Type}}(U \sigma)) \Rightarrow \text{RED}_{U \sigma} (s \sigma)
\]

Once it is proved, since we have \( \text{RED}_{\Delta_0}(\gamma) \), then by taking \( s = s_0, \Theta = \Delta_0, U = U_f \) and \( \sigma = \gamma \) we obtain the expected result.

**Remaining goal 2.** Let firstly justify that \( \text{RED}_{\Delta_0}(\gamma) \).

**Proof.** To prove it, we will have to perform an induction on the number of constructor under which is each variable. Let’s take an \( i \) between 1 and \( m \).

If \( p_i \) is a variable, we have by hypothesis that \( \text{RED} \left( T_f, i \left[ t_1/x_1 \ldots t_{i-1}/x_{i-1} \right] \right) (t_i) \).

And \( t_i = p_i \vartheta = \gamma(p_i) \).
Otherwise there is a $c_i$ such that $C(c_i) = \Pi(x_1 : U_1) \ldots (x_k : U_k) (d \tau_1 \ldots \tau_i)$ and $p_i = c_i p_{i,1} \ldots p_{i,k}$. We have by hypothesis that RED \((t_j,i)[t_{i,1}/x_1; \ldots; t_{i-1}/x_{i-1}]) (t_i)\). And $t_i = p_i \varrho = c_i (p_{i,1} \varrho) \ldots (p_{i,k} \varrho)$ and by definition of the reducibility of a term which begins with a constructor, we have RED \((U_j[p_{i,1}/x_1; \ldots; p_{i,j-1}/x_{j-1}]) (p_{i,j})\) and by induction, we can keep going under the constructors until we reach a variable.

**Remaining goal 3.** For all subterms $s$ of $s_0$, all contexts $\Theta$ extending $\Delta_0$, all $U$ and all substitutions $\sigma$ which are equal to $\gamma$ on $\Delta_0$, we have:

$$\Theta \vdash s : U \land \text{RED}_\Theta(\sigma) \land \text{RED}_{(\text{Type})}(U \sigma) \Rightarrow \text{RED}_{(U \sigma)}(s \sigma)$$

**Proof.** We will show it by induction on the subterms of $s_0$, knowing that the subterm-order is well-founded.

Let’s take a subterm $s$ of $s_0$, a context $\Theta$ extending $\Delta_0$, a $U$ and a substitution $\sigma$ which is equal to $\gamma$ on $\Delta_0$, such that $\Theta \vdash s : U$, RED_{(Type)}(U \sigma) and RED_{\Theta}(\sigma). We will proceed by case analysis on the form of $s$. Each form induces the end of the derivation of $\Theta \vdash s : U$. Here for readability, some hypothesis, which are necessary to really perform the inferences but which are not relevant to perform the induction are omitted.

$$\Theta \vdash s : T_i \begin{cases} F(g) = \Pi(x_1 : T_1).U_{\text{ar}(g)+1} \\
\text{If } n > \text{ar}(g), U_{\text{ar}(g)+1} \left[ s_1/x_1; \ldots, s_{\text{ar}(g)}/x_{\text{ar}(g)} \right] \Rightarrow \ast \Pi(x_{\text{ar}(g)+1} : T_{\text{ar}(g)+1}).U_{\text{ar}(g)+2} \\
\forall \text{ar}(g) + 1 < i \leq n. U_i \left[ s_{i-1}/x_{i-1} \right] \Rightarrow \ast \Pi(x_i : T_i).U_{i+1} \\
U \Rightarrow \ast U_{n+1} \left[ s_n/x_n \right] \end{cases}$$

By induction, we have RED_{(T_i \sigma)}(s_i \sigma). Since $g \ s_1 \ldots s_n$ is a subterm of $s_0$ and $\text{ar}(g) \geq n$, we have $(f, \bar{\sigma}) \succ (g, \bar{s_i})$. Here we must recall that we are under the hypothesis that

$$\forall i. \text{RED} \left( t_{j,i} \left[ t_{i,1}/x_1; \ldots; t_{i-1}/x_{i-1} \right] \right) (t_i) \text{ and } \forall g. \forall \bar{u}, (f, \bar{\sigma}) \succ (g, \bar{u}) \Rightarrow \Phi(g, \bar{u})$$

Since $\sigma$ and $\gamma$ are equal on $\Delta_0$ which is a context in which all the variables appearing in the $p_i$ are bounded, from $t_i \Rightarrow^* p_i \gamma$, we can deduce $t_i \Rightarrow^* p_i \sigma$. Furthermore, by proposition 2.3.10 we have WN($t_i$), so we get

$$(f, \bar{\sigma}) \succ (g, \bar{s_i} \bar{\sigma})$$

Then we get $\Phi(g, \bar{s_i} \bar{\sigma})$, hence RED \((U_{\text{ar}(g)+1}[\bar{s_i}/x_1]) ((g \ s_1 \ldots s_{\text{ar}(g)}) \sigma)\), then for all $j \geq \text{ar}(g)$, RED \((U_{j+1}[\bar{s_i}/x_1]) ((g \ s_1 \ldots s_j) \sigma)\) and by the conversion rule, RED_{(U \sigma)}(s \sigma).

$$\Theta \vdash s : T_i \left[ s_1/x_1, \ldots, s_{i-1}/x_{i-1} \right] \quad \Theta \vdash d \ s_1 \ldots s_k : \text{Type} \quad D(s) = \Pi(x_i : T_i) \text{ Type}$$
By induction, we have RED$(T_i, [s_1/x_1, \ldots, s_{i-1}/x_{i-1}] \sigma)(s_i \sigma)$. Then RED$_{\text{Type}}((d \sigma s_1 \ldots s_k) \sigma)$ which is our goal.

- $\Theta \vdash s_i : T_i \left[ s_1/x_1, \ldots, s_{i-1}/x_{i-1} \right] \\
\Theta \vdash \Gamma(s_1 \ldots s_m : U)$

By induction, we have RED$(T_i, [s_1/x_1, \ldots, s_{i-1}/x_{i-1}] \sigma)(s_i \sigma)$.

Then RED$_{\text{Type}}(d \sigma s_1 \ldots s_k) \left( \Gamma(s_1 \ldots s_m) \sigma \right)$ which by the conversion rule is RED$_{U \sigma}(s \sigma)$.

- $\Theta \vdash t : \Pi(x : T_1).T_2 \\
\Theta \vdash u : T_1 \\
\Theta \vdash tu : U$

By induction, we have RED$_{(U \sigma) T_1 \sigma}(u \sigma)$ and RED$_{\Pi(x : T_1).T_2 \sigma}(t \sigma)$, hence by definition we have RED$_{U \sigma}(t u \sigma)$.

- $\Theta, x : T \vdash s_1 : V \\
\Theta \vdash \lambda(x : T).s_1 : \Pi(x : T) V$

By hypothesis, we have RED$_{\text{Sort}}((\Pi(x : T) V) \sigma)$. Let’s take a $t$ such that RED$_{(\Pi \sigma) T \sigma}(t)$. Then by definition RED$_{\text{Sort}}(V [\sigma, t/\gamma]) \Theta, x : T$ enrich $\Delta_0$ because $\Theta$ do so and $[\sigma, t/\gamma]$ is equal to $\gamma$ on $\Delta_0$ because $\sigma$ is. So by induction, we have RED$_{\text{Type}}(V [\sigma, t/\gamma]) (s_1 [\sigma, t/\gamma]).$ Knowing that $((\lambda(x : T).s_1) \sigma t) \mapsto s_1 [\sigma, t/\gamma]$, we have RED$_{\text{Type}}((\lambda(x : T).s_1) \sigma t)$ and since $t$ was arbitrary, RED$_{\Pi(x : T) V \sigma}(\lambda(x : T).s_1) \sigma$ which was our goal.

By combining theorem 2.4.3, theorem 2.3.16 and proposition 2.3.10 we obtain the following corollary, which was the aim of our work, because it gives conditions under which typability implies weak normalisation.

**Corollary 2.4.4.** If $\prec$ is well-founded and RED$(\mathcal{F})$ then

$$\vdash t : T \Rightarrow \text{WN}(t)$$

Once we have this result, the *Size-Change Principle* can be use to verify if $\prec$ is well-founded. Thanks to the modularity of Wahlstedt’s work, the demonstration given in [Wah07, Chapter 4] can be re-used to prove this.

Inspired by the work of Frédéric Blanqui [Bla01], an adaptation of the reducibility predicates in order to show strong normalisation has been tried during the internship. We manage to show the theorem 2.3.16 with a reducibility predicate adapted for strong normalisation. Unhappily, we did not had the time to adapt the proof of theorem 2.4.3 for this predicate, if it is possible, and we did not used the $\lambda$II-calculus modulo theory, but the Wahlstedt’s system to do this adaptation. This is the reason why, this on-going work is not presented in this report, since it would have needed to introduce another system, which is not the
one in which our problem was designed, making this report much heavier and probably less clear, for a not finished work. We can expect this new adaptation to be foundable in another document soon, once it will be achieved.
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<td>$\Lambda$</td>
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<td>$l \rightarrow r$</td>
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<td>$\mathcal{N}$</td>
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