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ISSN 0926-4515

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Abstract

In this paper, an extension of Pure Type Systems (PTS's) to include definitions is presented and the meta-theory of these PTS's with definitions is treated in detail. We prove that all the properties of a PTS are preserved by the extension.
Chapter 1

Introduction

In this paper, the definitions and notations of [Bar92] will be used. In particular, we assume that the reader is familiar with the concepts of λ-cube and Pure Type Systems (PTS's) as defined in [Bar92]. The λ-cube forms a natural framework in which several systems à la Church are given in an uniform way. The description method of the systems in the λ-cube is generalised in the definition of a PTS. A PTS does not provide the possibility to introduce a definition, i.e. an abbreviation (name) for a larger term which can be used several times in a program or proof. In this paper, an extension of the PTS that includes definitions will be introduced.

Definitions are names abbreviating expressions. In \( (x = a : A \text{ in } b) \) the identifier \( x \) abbreviates the expression \( a \) of type \( A \) in the expression \( b \). We assume that \( x \) does not occur either in \( a \) or in \( A \). In other words, it is not a recursive definition.

Example 1.1. \( \text{id} = (\lambda x : \alpha. x) : \alpha \rightarrow \alpha \text{ in } (\lambda y : \alpha \rightarrow \alpha. \text{id}) \text{id}. \) The name \( \text{id} \) is an abbreviation for a larger term, i.e. \( \lambda y : \alpha \rightarrow \alpha. x \) and is used several times in the expression \( (\lambda y : \alpha \rightarrow \alpha. \text{id}) \text{id}. \)

Example 1.2.

1. Given \( \alpha, \beta : \ast \):

\[ \Lambda = (\lambda \alpha : \ast. \lambda \beta : \ast. (\Pi \gamma : \ast. (\alpha \rightarrow \beta \rightarrow \gamma)); (\ast \rightarrow \ast \rightarrow \ast) \text{ in } (\alpha \land \beta) \rightarrow \alpha \]

2. Given \( \alpha : \ast \):

\[ \perp = (\Pi \alpha : \ast. \alpha) \ast \text{ in } (\perp = (\lambda \alpha : \ast. \alpha \rightarrow \perp) \ast \rightarrow \ast \in (\alpha \rightarrow \perp)) \]

The intended meaning of a definition is that the definiendum \( x \) can be substituted by the definiens \( a \) in the expression \( b \). A definition \( (x = a : A \text{ in } b) \) can be considered as having a similar behaviour than \( (\lambda x : A. b)a \), i.e. the substitution of the variable \( x \) by \( a \) in the expression \( b \). In the β-reduction where \( (\lambda x : A. b)a \) reduces to \( b[x := a] \), the operation \( b[x := a] \) is the substitution of all the occurrences of \( x \) by \( a \) in the expression \( b \). In contrast to β-reduction, the expression \( (x = a : A \text{ in } b) \) reduces to an expression which has at most one occurrence of \( x \) unfolded by \( a \). In order to unfold definitions, a new relation called δ-reduction is introduced. We will show that a definition \( (x = a : A \text{ in } b) \) has the same behaviour as \( (\lambda x : A. b)a \) in the sense that the definition \( (x = a : A \text{ in } b) \) δ-reduces in several steps to \( b[x := a] \). Then the unfolding of definitions that is made via the δ-reduction corresponds to the operation of substitution \( b[x := a] \).

A definition \( (x = a : A \text{ in } b) \) is not another way of writing \( (\lambda x : A. b)a \). There are important differences between \( (x = a : A \text{ in } b) \) and \( (\lambda x : A. b)a \), both regarding their reduction behaviour and their typing. One of the reasons for considering \( (x = a : A \text{ in } b) \) and not \( (\lambda x : A. b)a \) is that in some cases it is convenient to have the freedom of substituting only in some of the occurrences of
an expression in a given formula.

Another reason for considering \((x = a : A \text{ in } b)\) and not \((\lambda x : A \cdot b)\) is that in \((x = a : A \text{ in } b)\) the fact that \(x\) is an abbreviation for \(a\) can be used to type \(b\). This is shown in the following example.

**Example 1.3.** The term \(\lambda x :* \cdot (x = a :* \text{ in } \lambda y : x \cdot \lambda f : a \rightarrow a \cdot f y)\) is typable in the system \(\lambda 2\) extended with definitions. But it is not possible to type the corresponding term expressed with an application and an abstraction in any system of the \(\lambda\)-cube,

\[
\lambda x :* \cdot (\lambda x :* \cdot \lambda y : x \cdot \lambda f : a \rightarrow a \cdot f y) a
\]

A third reason for considering \((x = a : A \text{ in } b)\) instead of \((\lambda x : A \cdot b)\) is that the abstraction \((\lambda x : A \cdot b)\) may not be allowed when \((x = a : A \text{ in } b)\) is.

**Example 1.4.** The term \((x = a :* \text{ in } \lambda y : x \cdot \lambda f : x \rightarrow z \cdot f y)\) is typable in the system \(\lambda\_\_\) extended with definitions. The corresponding term expressed with an application and an abstraction, i.e. 

\[(\lambda x :* \cdot \lambda y : x \cdot \lambda f : x \rightarrow z \cdot f y) a\]

is not typable in \(\lambda\_\_\) but it can be typable in \(\lambda 2\).

The fact that the type \(A\) of \(a\) is explicitly written in \((x = a : A \text{ in } b)\) is not essential, at least for PTS's where types are unique (up to \(\beta\)-equality). We want to record this type \(A\) in the case that we consider \(A\) as a proposition.

---

We present a short description of the paper.

In Chapter 2, we recall the definition of Pure Type Systems (PTS's) and their main properties from [Bar92].

In Chapter 3, Pure Type Systems with definitions (DPTS's) are introduced. The set of terms and contexts are extended to include definitions. A new relation \(\rightarrow\_\_\) is introduced that corresponds to the unfolding of a definition.

In Chapter 4, Church-Rosser property for \(\rightarrow\_\_\) and strong normalization property for \(\rightarrow\) are proved. All the properties in this chapter, in particular strong normalization for \(\rightarrow\), hold for all the pseudoterm systems, not only for well-typed terms.

In Chapter 5, the meta-theory of DPTS's is treated. All properties of PTS's with the exception of strong normalization are preserved by the extension. We prove that strong normalization is preserved for some PTS's such as the examples of PTS's given in [Bar92].

In Appendix A, a summary of the definitions of DPTS's is presented.

In Appendix B, the main features of the proofs of Church-Rosser for \(\delta\) and \(\beta\delta\)-reductions, weak normalization for \(\rightarrow\) and strong normalization for \(\rightarrow\_\_\) are abstracted and presented for Abstract Reduction Systems.
Chapter 2

Pure Type Systems

In this chapter the concept of Pure Type System (PTS) is defined as in [Bar92]. Only the notion of topsort and the system \( \lambda C \alpha \) (see definitions 2.11 and 2.14) are not defined in [Bar92].

**Definition 2.1.** The set \( T \) of pseudoterms and the set \( C \) of contexts are defined as follows:

\[
T ::= V \mid C \mid (T \cdot T) \mid (\lambda V : T \cdot T) \mid (\Pi V : T \cdot T)
\]

\[
C ::= \epsilon \mid < C, V : T >
\]

where \( V \) is the set of variables and \( C \) is the set of constants.

**Convention 2.2.** The following convention will be used in this paper:

1. Variables will be denoted as \( x, y, z, \ldots, \alpha, \beta, \gamma \ldots \)
   Contexts will be denoted as \( \Gamma, \Gamma', \Delta, \Delta', \ldots \)
   Terms will be denoted as \( a, b, d, \ldots, A, B, C, \ldots \)
   Constants will be denoted as \( c, c', \ldots \)
   The usual parenthesis conventions for abstraction, application and product will be used (see [Bar92]).

2. The usual parenthesis conventions for abstraction, application and product will be used (see [Bar92]).

**Definition 2.3.** Let \( a, a' \in T \). The concept "the term \( a \) \( \alpha \)-reduces to the term \( a' \)" is written as \( a \mapsto_\alpha a' \) and is defined by \( (\lambda x : A \cdot a)b \mapsto_\alpha a[x := b] \) and the compatibility rules.

The \( \alpha \)-equality is defined as usual. \( \alpha \)-equal terms are identified.

**Definition 2.4.** The specification of Pure Type System (PTS) is a triple \( S = (S, A, R) \) such that

- \( S \subseteq C \) is the set of sorts.
- \( A \subseteq C \times S \) is the set of axioms
- \( R \subseteq S \times S \times S \) is the set of rules

**Definition 2.5.** The PTS determined by the specification \( S = (S, A, R) \) is denoted as \( \lambda S = \lambda (S, A, R) \) and defined by the notion of type derivation \( \Gamma \vdash_{\lambda S} b : B \) (or \( \Gamma \vdash b : B \)) given by the following axioms and rules:

(axiom) \[
\Gamma \vdash c : s
\]

(start) \[
\Gamma \vdash A : s
\]

where \( x \) is \( \Gamma \)-fresh

for \( c : s \in A \)
Theorem 2.6. (Church Rosser for $\beta$-reduction) Let $\Gamma \in C$ and $a \in T$ be such that $a \rightarrow_{\beta} b$ and $a \rightarrow_{\beta} c$. Then there exists a term $d \in T$ such that $b \rightarrow_{\beta} d$ and $c \rightarrow_{\beta} d$.

The following properties are in [Bar92].

Theorem 2.7. (Correctness of Types) Let $\Gamma \in C$ and $d, d', D \in T$ be such that $\Gamma \vdash d : D$. Then $\Gamma \vdash D : s$ or $D = s$.

Theorem 2.8. (Subject Reduction Theorem) Let $\Gamma \in C$ and $d, d', D \in T$ be such that $\Gamma \vdash d : D$. If $d \rightarrow_{\beta} d'$ then $\Gamma \vdash d' : D$.

Definition 2.9. The specification $S = (S, A, R)$ is called singly sorted if

1. $(c : s_1), (c : s_2) \in A$ implies $s_1 \equiv s_2$

2. $(s_1, s_2, s_3), (s_1, s_2, s_3) \in R$ implies $s_3 \equiv s_3$

Theorem 2.10. (Uniqueness of Types) Let $S$ be a singly sorted specification, $\Gamma \in C$ and $a, A, B \in T$ such that $\Gamma \vdash a : A$ and $\Gamma \vdash a : B$. Then $A \equiv B$.

Definition 2.11. Let $\lambda S$ be a PTS. A sort $s$ in $S$ is called a topsort if there is no $s_0 \in S$ such that $\Gamma \vdash s : s_0$.

Definition 2.12. The $\lambda$-cube is a cube of eight systems defined by the same set of sorts $S = \{*, \square\}$ and the same set of axioms $A = \{*, \square\}$. They differ in the set of rules $R$.

<table>
<thead>
<tr>
<th>System</th>
<th>$R$</th>
</tr>
</thead>
<tbody>
<tr>
<td>$\lambda_1$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda_2$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda P$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda P_2$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda \omega$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda P \omega$</td>
<td>$(*, *)$</td>
</tr>
<tr>
<td>$\lambda P \omega \equiv \lambda C$</td>
<td>$(*, *)$</td>
</tr>
</tbody>
</table>

The rule $(s_1, s_2)$ is an abbreviation for $(s_1, s_2, s_3)$.

The system $\lambda C$ is the Calculus of Constructions.
There is only one topsort in the $\lambda$-cube and that is $\square$. 
Definition 2.13. The system of higher order logic can be described by the following PTS:

\[ \lambda \text{HOL} \]
\[ S \quad \ast, \square, \Delta \]
\[ A \quad \ast : \square, \square : \Delta \]
\[ R \quad (\ast, \ast), (\square, \ast), (\square, \square) \]

There is only one topsort in \( \lambda \text{HOL} \) and that is \( \Delta \).

Definition 2.14. The Calculus of Constructions extended with an infinite type hierarchy can be described by the following PTS:

\[ \lambda C_{\infty} \]
\[ S \quad \in \]
\[ A \quad \forall n \in \mathbb{N} \quad n : n + 1 \]
\[ R \quad \forall m \in \mathbb{N} \quad (m, 0, 0) \&
\[ \forall m, n \in \mathbb{N} \quad \max(m, n) \leq r \quad (m, n, r) \]

The system \( \lambda C_{\infty} \) extended with strong \( \Sigma \)-types and cumulativity is the system ECC (see [Luo]). We can see that \( \lambda C_{\infty} \) is an extension of \( \lambda C \) and of \( \lambda \text{HOL} \) writing \( \ast \) instead of 0, \( \square \) instead of 1 and \( \Delta \) instead of 3.

Note that there is no topsort in \( \lambda C_{\infty} \).

Definition 2.15. Let \( AS \) be a PTS. Then \( AS \) is \( \beta \)-strongly normalizing if \( a \) and \( A \) \( \beta \)-strongly normalize for all \( a, A \in T \) and \( \Gamma \in C \) such that \( \Gamma \vdash_{AS} a : A \).

Theorem 2.16.

1. The system \( \lambda C_{\infty} \) is \( \beta \)-strongly normalizing.
2. The systems of the \( \lambda \)-cube are \( \beta \)-strongly normalizing.
3. The system \( \lambda \text{HOL} \) is \( \beta \)-strongly normalizing.

Proof: The system ECC is strongly normalizing (see [Luo]) and contains \( \lambda C_{\infty} \), the systems of the \( \lambda \)-cube and \( \lambda \text{HOL} \).

Not all PTS's are strongly normalizing.

Example 2.17. The PTS \( \lambda \ast \) determined by the specification \( (S, A, R) \) where \( S = \{ \ast \} \), \( A = \{ \ast : \ast \} \) and \( R = \{ (\ast, \ast) \} \) is not \( \beta \)-strongly normalizing.
Chapter 3

Pure Type Systems with definitions

In Section 3.1, the set of pseudoterms $T$ and the set of contexts $C$ defined in chapter 2 are extended to include definitions.

In Section 3.2, a new relation $\rightarrow_d$ between pseudoterms that corresponds to the unfolding of a definition is introduced. Also, the relation $\rightarrow_D$ is extended to the new pseudoterms.

In Section 3.3, Pure Type Systems with definitions (DPTS's) are introduced by extending PTS's with the new pseudoterms and contexts.

3.1 Pseudoterms

Definition 3.1.1. The set $T_\delta$ of pseudoterms is given by

$$\begin{align*}
T_\delta & := V \mid C \mid (T_\delta T_\delta) \mid (\lambda V:T_\delta T_\delta) \mid (\Pi V:T_\delta T_\delta) \mid (V=T_\delta T_\delta \text{ in } T_\delta)
\end{align*}$$

where $V$ is the set of variables and $C$ is the set of constants.

The mapping $FV : T \rightarrow P(V)$ defined in [Bar92] is extended to $T_\delta$.

Definition 3.1.2. The mapping $FV : T_\delta \rightarrow P(V)$ is defined as follows:

$$\begin{align*}
FV(x) &= \{x\} \\
FV(c) &= \emptyset \\
FV(a \cdot b) &= FV(a) \cup FV(b) \\
FV(\lambda x:A. a) &= FV(A) \cup (FV(a) - \{x\}) \\
FV(\Pi x:A. a) &= FV(A) \cup (FV(a) - \{x\}) \\
FV(x = a:A \text{ in } b) &= FV(A) \cup FV(a) \cup (FV(b) - \{x\})
\end{align*}$$

Definition 3.1.3. The set $C_\delta$ of contexts is given by

i) $\epsilon \in C_\delta$

ii) $\langle \Gamma, x:A \rangle \in C_\delta$ if $\Gamma \in C_\delta$, $x \in V$, $A \in T_\delta$ and $x$ is $\Gamma$-fresh

iii) $\langle \Gamma, z = a:A \rangle \in C_\delta$ if $\Gamma \in C_\delta$, $z \in V$, $a \in T_\delta$, $A \in T_\delta$, $z$ is $\Gamma$-fresh and $z \notin FV(a) \cup FV(A)$

Note that the set of contexts is not given by $C_\delta := \epsilon \mid C_\delta, V : T_\delta \mid C_\delta, V = T_\delta : T_\delta$. In order to avoid capture of bound variables in the definition of $\delta$-reduction given below a new variable is introduced to a context $\Gamma$ only if it is $\Gamma$-fresh, i.e. it does not occur in $\Gamma$.

It is said that the variable $z$ abbreviates the expression $a$ in $(z = a:A \text{ in } b)$ and in $\Gamma, z = a:A$.

Convention 3.1.4. The following convention about contexts will be used in this paper:
1. The expression $\Gamma, x:A$ stands for the context $<\Gamma, x:A>$.

2. The expression $\Gamma, x=a:A$ stands for the context $<\Gamma, x=a:A>$.

**Definition 3.1.5.** A mapping $Dom: C \rightarrow P(V)$ is defined as follows:

- $Dom(c) = \emptyset$
- $Dom(\Gamma, x:A) = Dom(\Gamma) \cup \{x\}$
- $Dom(\Gamma, x=a:A) = Dom(\Gamma) \cup \{x\}$

**Definition 3.1.6.** A mapping $Def: C \rightarrow P(V)$ is defined as follows:

- $Def(c) = \emptyset$
- $Def(\Gamma, x:A) = Def(\Gamma)$
- $Def(\Gamma, x=a:A) = Def(\Gamma) \cup \{x\}$

Substitution is extended to pseudoterms $T$ and to contexts $C$.

**Definition 3.1.7.** Let $M, N \in T$, $x \in V$. The result of substituting $N$ for (the free occurrences of) a variable $x$ in $M$ is denoted as $M[x := N]$ and is defined as follows:

$$
\begin{align*}
  c[x := N] &\equiv c & c \in C \\
  x[x := N] &\equiv N \\
  y[x := N] &\equiv y \\
  (\lambda x:A. a)[x := N] &\equiv (\lambda x:A. a) \\
  (\lambda y:A. a)[x := N] &\equiv (\lambda y:A[x := N]. a[x := N]) & y \notin FV(N) \\
  (\lambda y:A. a)[x := N] &\equiv (\lambda y:A[y := z][x := N]. a[y := z][x := N]) & y \notin FV(N) \\
  (a \ b)[x := N] &\equiv (a[x := N] \ b[x := N]) \\
  (\Pi x:A. a)[x := N] &\equiv (\Pi x:A. a) \\
  (\Pi y:A. a)[x := N] &\equiv (\Pi y:A[x := N]. a[x := N]) & y \notin FV(N) \\
  (\Pi y:A. a)[x := N] &\equiv (\Pi y:A[y := z][x := N]. a[y := z][x := N]) & y \notin FV(N) \\
  (x=a:A \ in \ b)[x := N] &\equiv (x=a:A \ in \ b) \\
  (y=a:A \ in \ b)[x := N] &\equiv (y=a[x := N]:A[x := N] \ in \ b[x := N]) & y \notin FV(N) \\
  (y=a:A \ in \ b)[x := N] &\equiv (z=a[x := N]:A[x := N] \ in \ b[y := z][x := N]) & y \notin FV(N)
\end{align*}
$$

**Definition 3.1.8.** The result of substituting $N$ for (the free occurrences of) a variable $x$ in $\Gamma$ such that $x \notin Dom(\Gamma)$ is denoted as $\Gamma[x := N]$ and is defined as follows:

$$
\begin{align*}
  c[x := N] &\equiv c \\
  x[x := N] &\equiv N \\
  y[x := N] &\equiv y
\end{align*}
$$

The definitions of change of bound variables and $\alpha$-congruence in $\lambda$-terms are extended to pseudoterms $T$.

**Definition 3.1.9.** Let $t \in T$, be such that $t$ contains the term $(x=a:A \ in \ b)$. A change of a bound variable in the term $t$ is the replacement of a subterm $(x=a:A \ in \ b)$, $(\lambda x:A. b)$ or $(\Pi y:A. b)$ by $(y=a:A \ in \ b[y := y])$, $(\lambda x:A. b[x := y])$ or $(\Pi y:A. b[x := y])$, respectively, where $y \notin FV(b)$.

The relation of $\alpha$-equality between $\lambda$-terms is extended to pseudoterms $T$.

**Definition 3.1.10.** Let be $\Gamma \in C$ and $t, t' \in T$, the term $t$ is $\alpha$-equal to $t'$ (written $t \equiv_{\alpha} t'$) if $t'$ is the result of applying to $t$ a series of changes of variables or vice versa.

**Convention 3.1.11.** Two terms are identified if they are $\alpha$-equal.  

---

1. In this definition the variable $y$ is different from $x$ and $z$ is fresh.
3.2 Reductions

The relation $\rightarrow_{\beta}$ between terms in $T$ will be extended to terms in $T_{6}$.

Definition 3.2.1. Let $a, a' \in T_{6}$. The concept "the term $a$ $\beta$-reduces to the term $a'$" is written as $a \rightarrow_{\beta} a'$ and is defined as usual by the following rules:

\[
\begin{align*}
(\lambda x: A. \ a)b & \rightarrow_{\beta} a[x := b] \\
(a \ b) & \rightarrow_{\beta} (a' \ b) \\
(\Pi x: A. \ a) & \rightarrow_{\beta} (\Pi x: A'. \ a)
\end{align*}
\]

A new relation $\rightarrow_{\delta} \subseteq C_{5} \times T_{6} \times T_{6}$ is introduced.

Definition 3.2.2. Let $\Gamma \in C_{5}$, and $a, a' \in T_{6}$. The concept "the term $a$ $\delta$-reduces to the term $a'$ in the context $\Gamma$" is written as $\Gamma \vdash a \rightarrow_{\delta} a'$ and is defined by the following rules:

\[
\begin{align*}
\Gamma, x = a: A, \Gamma_{2} \vdash x \rightarrow_{\delta} a & \\
\Gamma \vdash (x = a: A \text{ in } b) \rightarrow_{\delta} b & \\
\Gamma, x = a: A \vdash b \rightarrow_{\delta} b' & \\
\Gamma \vdash (x = a: A \text{ in } b) \rightarrow_{\delta} (x = a: A \text{ in } b') & \\
\Gamma \vdash a \rightarrow_{\delta} a' & \\
\Gamma \vdash (a \ b) \rightarrow_{\delta} (a' \ b) & \\
\Gamma, x: A \vdash a \rightarrow_{\delta} a' & \\
\Gamma, x: A \vdash (\lambda x: A. \ a) \rightarrow_{\delta} (\lambda x: A'. \ a') & \\
\Gamma, x: A \vdash a \rightarrow_{\delta} a' & \\
\Gamma \vdash (\Pi x: A. \ a) \rightarrow_{\delta} (\Pi x: A'. \ a') & \\
\end{align*}
\]

Note that the first rule allows to unfold definitions, the second rule allows to remove definitions and the remaining are compatibility rules. In a context $\Gamma, x = a: A$ a term $d$ can be $\delta$-reduced to a term $d'$ where $d'$ is obtained replacing one occurrence of $x$ by $a$ in $d$.

When $\Gamma$ is the empty context, $a \rightarrow_{\delta} a'$ is written instead of $\Gamma \vdash a \rightarrow_{\delta} a'$. 

8
Example 3.2.3.

1. Assume \( \Gamma \equiv A \cdot \).

\[
\begin{align*}
\Gamma \vdash (id= (\lambda y. A. y)\text{ in } (\lambda x. (\Pi x. A. A). id) \quad id \rightarrow_8 \\
(id=(\lambda y. A. y)\text{ in } (\lambda x. (\Pi x. A. A). (\lambda y. A. y)) \quad id \rightarrow_8 \\
(id=(\lambda y. A. y) \text{ in } (\lambda x. (\Pi x. A. A). (\lambda y. A. y)) \quad (\lambda y. A. y) \rightarrow_8 \\
(\lambda x. (\Pi x. A. A). (\lambda y. A. y)) \quad (\lambda y. A. y)
\end{align*}
\]

2. Assume \( \Gamma \equiv A : *, B : *, C : *, \rightarrow = \lambda \alpha : * . \lambda \beta : * . (\Pi x. \alpha . \beta) : \Pi x. * . (\Pi \beta : * . *) \),

\[
\begin{align*}
K=\langle \lambda x. \lambda y.B. x \rangle; (A \rightarrow B \rightarrow A), \\
S=\langle \lambda x: (A \rightarrow B \rightarrow C). \lambda y: (A \rightarrow B). \lambda z: A. z \cdot (y \cdot z) \rangle; \\
((A \rightarrow B \rightarrow C) \rightarrow (A \rightarrow B) \rightarrow (A \rightarrow C)) >
\end{align*}
\]

\[A \rightarrow B\] is written instead of \((\rightarrow A)B\)

The examples above show that in each step of a \( \delta \)-reduction, exactly one occurrence of a variable is unfolded. The first example also shows that the definition is removed in the last step of \( \delta \)-reductions.

The next example shows that \( \alpha \)-conversion is necessary when rewriting terms.

Example 3.2.4.

\[
x=y:A \vdash (y=z:u \text{ in } x)
\]

\[=_\alpha (y'=z:u \text{ in } x)
\]

\[\rightarrow_8 (y'=z:u \text{ in } y)
\]

The variable \( y \) occurs in the term \((y=z:u \text{ in } x)\) and in the context \( x=y:A \).

Convention 3.2.5. From now on, the bound variables in the term will be assumed different from the variables in the context.

Definition 3.2.6. Let \( \Gamma \in C_\delta \) and \( a, a' \in T_\delta \). Then

\[\Gamma \vdash a \rightarrow_{\beta \delta} a' \text{ if } \Gamma \vdash a \rightarrow_8 a' \text{ or } a \rightarrow_\beta a'.\]

Definition 3.2.7. The relation \( \rightarrow^{=}_{\rho} \) is the reflexive closure of \( \rightarrow_\rho \). The relation \( \rightarrow^=_{\tau} \) is the transitive closure of the reduction relation \( \rightarrow_\tau \). The relation \( \rightarrow_{\rho} \) is the transitive, reflexive closure of the reduction relation \( \rightarrow_\rho \). The relation \( =_\rho \) is the equivalence relation generated by \( \rightarrow_{\rho} \).

Note that for \( R \in \{=_8, =_1, \rightarrow_{\beta \delta}, =_{\beta \delta} \} \) a context has to be specified, i.e. \( \Gamma \vdash a \rightarrow_{Rb} \) for \( a, b \in T_\delta \) and \( \Gamma \in C_\delta \).

3.3 Types

We extend the notion of Pure Type System to include definitions.

Definition 3.3.1. The DPTS determined by the specification \( S = (S, A, R) \) is denoted as \( \lambda S_\delta = \lambda (S, A, R)_\delta \) and defined by the notion of type derivation \( \Gamma \vdash_{\lambda S_\delta} b : B \) (or just \( \Gamma \vdash b : B \)) with \( \Gamma \in C_\delta \) and \( b, B \in T_\delta \) given by the following axioms and rules:
(axiom) \[ \varepsilon \vdash c : s \quad \text{for } c : s \in A \]

(start) \[ \Gamma \vdash A : s \]
\[ \Gamma, x : A \vdash x : A \]

(weakening) \[ \Gamma \vdash b : B \quad \Gamma \vdash A : s \]
\[ \Gamma, x : A \vdash b : B \]

(formation) \[ \Gamma \vdash A : s_1 \quad \Gamma, x : A \vdash B : s_2 \]
\[ \Gamma \vdash (\Pi x : A. B) : s_3 \]

(abstraction) \[ \Gamma, x : A \vdash b : B \quad \Gamma \vdash (\Pi x : A. B) : s \]
\[ \Gamma \vdash (\lambda x : A. b) : (\Pi x : A. B) \]

(application) \[ \Gamma \vdash b : (\Pi x : A. B) \quad \Gamma \vdash a : A \]
\[ \Gamma \vdash (b \ a) : B[x := a] \]

(conversion) \[ \Gamma \vdash b : B \quad \Gamma \vdash B' : s \quad B = B' \]
\[ \Gamma \vdash b : B' \]

(\(\delta\) - start) \[ \Gamma \vdash a : A \]
\[ \Gamma, x = a : A \vdash x : A \]

(\(\delta\) - weakening) \[ \Gamma \vdash b : B \quad \Gamma \vdash a : A \]
\[ \Gamma, x = a : A \vdash b : B \]

(\(\delta\) - formation) \[ \Gamma, x = a : A \vdash B : s \]
\[ \Gamma, x = a : A \vdash b : B \]
\[ \Gamma \vdash (x = a : A \text{ in } B) : s \]

(\(\delta\) - introduction) \[ \Gamma, x = a : A \vdash b : B \quad \Gamma \vdash (x = a : A \text{ in } B) : s \]
\[ \Gamma \vdash (x = a : A \text{ in } B) : (x = a : A \text{ in } B) \]

(\(\delta\) - conversion) \[ \Gamma \vdash b : B \quad \Gamma \vdash B' : s \quad \Gamma \vdash B = B' \]
\[ \Gamma \vdash b : B' \]

where \(s\) ranges over sorts, i.e. \(s \in S\).

The condition "\(x\) is \(\Gamma\)-fresh" in \(\delta\)-start and \(\delta\)-weakening rules could be omitted because of convention 3.1.11.

Observe that the new rules are: \(\delta\)-start, \(\delta\)-weakening, \(\delta\)-formation, \(\delta\)-introduction and \(\delta\)-conversion.

The system \(\lambda S_5\) is an extension of \(\lambda S\).

Observe that in a DPTS, only the typable terms can be abbreviated. This is because the \(\delta\)-start and \(\delta\)-weakening rules allow the abbreviation of the term \(a\) in the context \(\Gamma, x = a : A\) only if \(\Gamma \vdash a : A\). Hence it is not possible to abbreviate any topsort. For example in the \(\lambda\)-cube it is not possible to abbreviate \(\Box\).

Observe that if \(B\) is a toposort then \((x = a : A \text{ in } B)\) is neither typable nor the type of another definition \((x = a : A \text{ in } b)\). This is because the \(\delta\)-formation rule requires that \(\Gamma, x = a : A \vdash B : s\) and because the \(\delta\) introduction rule requires that \(\Gamma \vdash (x = a : A \text{ in } B) : s\). For example \((x = a : A \text{ in } \Box)\) is not typable in \(\lambda C_5\) and it is not a type of \((x = a : A \text{ in } * )\). The condition "\(\Gamma \vdash (x = a : A \text{ in } B) : s\)" in the premises of the \(\delta\)-introduction rule is necessary to ensure Subject Reduction Theorem. Since we add this condition we must also add \(\delta\)-formation rule. We call \((\lambda S)_5\) the system obtained from \(\lambda S_5\) by removing the \(\delta\)-formation rule and the condition "\(\Gamma \vdash (x = a : A \text{ in } B) : s\)" from the premises of the \(\delta\)-introduction rule. Subject Reduction Theorem does not hold in \((\lambda S)_5\) as the following example shows.

**Example 3.3.2.** Suppose \(\Gamma \vdash a : A\). The term \((x = a : A \text{ in } * )\) has type \((x = a : A \text{ in } \Box)\) in \((\lambda C)_5\). This term \(\delta\)-reduces to * but * does not have type \((x = a : A \text{ in } \Box)\). In order to derive \(\Gamma \vdash * : (x = a : A \text{ in } \Box)\) we apply \(\delta\)-conversion rule

\[
\Gamma \vdash * : \Box \quad \Gamma \vdash (x = a : A \text{ in } \Box) : s \quad \Gamma \vdash (x = a : A \text{ in } \Box) = s \Box
\]
\[ \Gamma \vdash * : (x = a : A \text{ in } \Box) \]
But \((z=a:A)\) in \(\Box\) is not typable because \(\Box\) is not.

**Example 3.3.3.**

- Let \(\Gamma_1 \equiv \alpha : *, ID = (\Pi x : \alpha . \alpha):*, id = (\lambda x : \alpha . x) : ID >\). In \((\lambda \to)_d\), the following can be derived.

  \[ \Gamma_1 \vdash id : ID \]

- Let \(\Gamma_2 \equiv \gamma : *, \neg = \lambda \alpha : *. \lambda \beta : *. (\Pi x : \alpha . \beta)(\Pi \alpha : *. \Pi \beta : *. *) \).

  In \(\lambda \Pi \delta\), the following can be derived.

  \[ \Gamma_2 \vdash \lambda x : \gamma. x : (\gamma \to \gamma) \]

- Let \(\Gamma_3 \equiv (\exists x : \alpha : *, \neg = (\lambda \alpha : *. \alpha \to \bot)(\Pi \alpha : *. *) \).

  In \(\lambda \psi\), the following can be derived.

  \[ \Gamma_3 \vdash (\lambda x : \gamma. \lambda y : \gamma. (y \ x)): (\gamma \to \gamma) \]

We can not abbreviate the expression \(\lambda \alpha : \Box. \lambda \beta : \Box. (\Pi x : \alpha . \beta)\) because it is not typable in the \(\lambda\)-cube.

- Let \(\Gamma_4 \equiv (\exists x, \alpha_1 : *, \alpha_2 : *, \neg = (\lambda \alpha : *. \lambda \beta : *.(\Pi \gamma : *(\alpha \to \beta \to \gamma)))(\Pi \alpha : *. \Pi \beta : *. *), K = (\lambda \alpha : *. \lambda \beta : *. \lambda \alpha : \lambda \gamma : \beta. x): (\Pi \alpha : *. \Pi \beta : *. \alpha \to \beta \to \alpha) \).

  In \(\lambda \psi\), the following can be derived.

  \[ \Gamma_4 \vdash (\lambda x : (\lambda \alpha_1 \alpha_2). x \alpha_1 (K \alpha_1 \alpha_2)) : (\lambda \alpha_1 \alpha_2) \to \alpha_1 \]

We will prove the following properties:

- Church Rosser Theorem for \(\to_{\beta}\) and for \(\to_{\beta \delta}\).

- Strong Normalization Theorem for \(\to_{\beta}\).

- Subject Reduction Theorem for DPTS's, i.e. if \(\Gamma \vdash d : D\) and \(\Gamma \vdash d \to_{\beta \delta} d'\) then \(\Gamma \vdash d' : D\).

- Uniqueness of Types for singly sorted DPTS's, i.e. if \(\Gamma \vdash d : D\) and \(\Gamma \vdash d : D'\) then \(D =_{\beta \delta} D'\).

- Conservativity, i.e. for \(A \in T, \Gamma \in C \exists a \Gamma \vdash_{AS} a : A\) iff \(\exists a \Gamma \vdash_{AS} a : A\)

- If a PTS is \(\beta \delta\)-weakly normalizing then the corresponding DPTS is \(\beta \delta\)-weakly normalizing too.

- An extension of a PTS is \(\beta \delta\)-strongly normalizing if a "slightly" larger PTS is \(\beta\)-strongly normalizing. In particular, the Calculus of Constructions extended with definitions is strongly normalizing.
Chapter 4

Properties of General Pseudoterms

In this chapter, basic properties of $\beta$ and $\delta$-reductions for all pseudoterms are proved.

In Section 4.2, Church Rosser property is proved for $\rightarrow_\delta$ and for $\rightarrow_\beta_\delta$.

The idea of the proof of Church Rosser for $\delta$-reduction is as follows:

- A "projection" map, $| - |_\Gamma : C \times T_\delta \rightarrow T$, is defined. The "projection" $|a|_\Gamma$ is a term that is obtained from $a$ by unfolding all the definitions occurring in $\Gamma$ and in $a$.

- A term $a$ $\delta$-reduces to its projection, i.e. $\Gamma \vdash a \rightarrow_\delta |a|_\Gamma$.

- The projections of two terms that are in $\rightarrow_\delta$ are equal, i.e. if $\Gamma \vdash c \rightarrow_\delta d$ then $|c|_\Gamma = |d|_\Gamma$.

Finally, the proof of Church Rosser for $\beta\delta$-reduction follows from the previous considerations and the following one:

- The projection preserves $\beta$-reduction, i.e. if $\Gamma \vdash c \rightarrow_\delta d$ then $|c|_\Gamma \rightarrow_\beta |d|_\Gamma$.

In Section 4.3, Strong Normalization property is proved for $\rightarrow_\delta$.

Also, an illustrative and intuitive proof of Weak Normalization for $\delta$-reduction is presented. The $\delta$-normal form of a term $a$ in a context $\Gamma$ is its projection, i.e. the result from $a$ by unfolding all the definitions occurring in $\Gamma$ and in $a$.

In order to prove Strong Normalization for $\rightarrow_\delta$, the well-known method of defining a function $\text{nat}_{\rightarrow_\delta} : C \times T_\delta \rightarrow \mathbb{N}$ which decreases with $\delta$-reduction is used.

4.1 Basic Properties

Lemma 4.1.1. Let $\Gamma_1, \Gamma_2, \Gamma_3 \in C_A$ and $b, b' \in T_\delta$ be such that $\Gamma_1, \Gamma_3 \vdash b \rightarrow_\delta b'$. Then

$$\Gamma_1, \Gamma_2, \Gamma_3 \vdash b \rightarrow_\delta b'.$$

Proof: It is proved by induction on the definition of $\rightarrow_\delta$.

Lemma 4.1.2. Let be $\Gamma \in C_A$, $x \in V$ and $a, A, b, b' \in T_\delta$. The following rule is derived from the ones in the definition of $\rightarrow_\delta$.

$$\Gamma \vdash b \rightarrow_\delta b' \quad \Gamma \vdash (x = a : A \text{ in } b) \rightarrow_\delta (x = a : A \text{ in } b')$$

Proof: By lemma 4.1.1, it follows that $\Gamma, x = a : A \vdash b \rightarrow_\delta b'$. By definition of $\rightarrow_\delta$, it follows that $\Gamma \vdash (x = a : A \text{ in } b) \rightarrow_\delta (x = a : A \text{ in } b')$. 

\[12\]
Lemma 4.1.3.

1. Let \(<\Gamma, x:A, \Gamma'>\in \mathcal{C}_t\) and \(b\in \mathcal{T}_t\).
   \(\Gamma, x:A, \Gamma' \vdash b \rightarrow_{\beta_S} b'\) if and only if \(\Gamma, \Gamma' \vdash b \rightarrow_{\beta_S} b'\).

2. Let \(<\Gamma, x=a:A, \Gamma'>\in \mathcal{C}_t\) and \(b\in \mathcal{T}_t\) be such that \(x \notin \text{FV}(b)\).
   \(\Gamma, x=a:A, \Gamma' \vdash b \rightarrow_{\beta_S} b'\) if and only if \(\Gamma, \Gamma' \vdash b \rightarrow_{\beta_S} b'\).

Proof:

1. The implication from left to right is proved by induction on the definition of \(\rightarrow_{\beta_S}\). The converse implication follows by lemma 4.1.1.

2. The implication from left to right is proved by induction on the definition of \(\rightarrow_{\beta_S}\). The converse implication follows by lemma 4.1.1.

Example 4.1.4. In example 3.2.3 we show that \(id=(\lambda y:A. y):(\Pi y:A. A)\) in \((\lambda x:(\Pi y:A. A). id)id\) reduces in several steps to the term

\((\lambda x:(\Pi y:A. A). (\lambda y:A. y))(\lambda y:A. y)\)

Note that \((\lambda x:(\Pi y:A. A). (\lambda y:A. y))(\lambda y:A. y)\) is the substitution of all the occurrences of \(id\) by \((\lambda y:A. y)\) in the expression \((\lambda x:(\Pi y:A. A). id)id\), i.e.

\((\lambda x:(\Pi y:A. A). id)|id|\equiv (\lambda x:(\Pi y:A. A). (\lambda y:A. y))(\lambda y:A. y)\)

This is a particular case of the following theorem:

Theorem 4.1.5. Let \(\Gamma \equiv <\Gamma_1, x=a:A, \Gamma_2>\in \mathcal{C}_t\) and \(b\in \mathcal{T}_t\). Then

\(\Gamma \vdash b \rightarrow_S b[x := a]\).

Proof: It follows easily by induction on the structure of \(b\). \(\square\)

Corollary 4.1.6. Let \(\Gamma \in \mathcal{C}_t\) and \((x=a:A in b)\in \mathcal{T}_t\). Then

\(\Gamma \vdash (x=a:A in b) \rightarrow_S b[x := a]\)

Proof: By theorem 4.1.5, it follows that \(\Gamma, x=a:A \vdash b \rightarrow_S b[x := a]\). Then \(\Gamma \vdash (x=a:A in b) \rightarrow_S (x=a:A in b[x := a])\). Since \(x \notin \text{FV}(b[x := a])\), it follows that \((x=a:A in b[x := a]) \rightarrow_S (x=a:A in b[x := a])\) \(\square\)

Lemma 4.1.7. [Substitution Lemma] Let \(a, b, d \in \mathcal{T}_t\). Suppose \(x \neq y\) and \(x \notin \text{FV}(d)\). Then

\(a[x := b][y := d] \equiv a[y := d][x := b[y := d]]\)

Proof: It is proved by induction on the structure of \(a\). \(\square\)

Lemma 4.1.8. Given \(a, b \in \mathcal{T}_t\) such that \(a \rightarrow_S a'\). Then

\(a[x := b] \rightarrow_S a'[x := b]\)

Proof: It is proved by induction on the generation of \(\rightarrow_S\). \(\square\)

The previous lemma means that \(\beta\) is substitutive. The following example shows that \(\rightarrow_S\) is not substitutive.
Example 4.1.9. Let $\Gamma$ be the context $< A : \ast, id=(\lambda y : A. y)(\Pi y : A. A) >$.

$\Gamma \vdash id \rightarrow_\delta (\lambda y : A. y)$

But it is not true that $\Gamma \vdash id[ id := b ] \rightarrow_\delta (\lambda y : A. y)[id := b]$ for all $b \in T_6$. This is because $id \in Def(\Gamma)$.

Lemma 4.1.10. Given $a, b \in T_6$ such that $\Gamma \vdash b \rightarrow_\delta b'$. Then

$\Gamma \vdash a[x := b] \rightarrow_\delta a[x := b']$.

Proof: It is proved by induction on the structure of $a$.

- Assume $a$ is a variable. There are two possible cases.
  - Assume $a \equiv x$. By hypothesis, $\Gamma \vdash x[x := b] \rightarrow_\delta x[x := b']$.
  - Assume $a \equiv y$ and that $y$ is different from $x$. Hence $\Gamma \vdash y[x := b] \rightarrow_\delta y[x := b']$.

- Assume $a \equiv (y=c:D \in e)$. There are two possible cases.
  - Assume $y \equiv x$. Hence $\Gamma \vdash (y=c:D \in e)[x := b] \rightarrow_\delta (y=c:D \in e)[x := b']$.
  - Assume $y$ is different from $x$. It follows from IH that $\Gamma \vdash c[x := b] \rightarrow_\delta c[x := b']$, $\Gamma \vdash D[x := b] \rightarrow_\delta D[x := b']$ and $\Gamma \vdash e[x := b] \rightarrow_\delta e[x := b']$. It follows from lemma 4.1.2, and the definition of $\rightarrow_\delta$ that $\Gamma \vdash (y=c[D := b])[x := b] \rightarrow_\delta (y=c[D := b'])[x := b]]$.

Then by definition of substitution $\Gamma \vdash (y=c:D \in e)[x := b] \rightarrow_\delta (y=c:D \in e)[x := b']$.

Theorem 4.1.11.

1. If $\Gamma \vdash e \rightarrow_\delta e'$ and $\Gamma, x=e:E, \Gamma' \vdash B \rightarrow_\delta B'$ then $\Gamma, x=e:E, \Gamma' \vdash B \rightarrow_\delta B'$.

2. If $\Gamma, x=e:E, \Gamma' \vdash B \rightarrow_\delta B'$ then $\Gamma, \Gamma'[x := e] \vdash B[x := e] \rightarrow_\delta B'[x := e]$.

Proof:

1. It is proved by induction on the definition of $\Gamma, x=e:E, \Gamma' \vdash B \rightarrow_\delta B'$.

2. It is proved that if $\Gamma, x=e:E, \Gamma' \vdash b \rightarrow_\delta b'$ then $\Gamma, \Gamma'[x := e] \vdash b[x := e] \rightarrow_\delta b'[x := e]$ (by induction on the definition of $\rightarrow_\delta$).

By lemma 4.1.8, it follows that if $\Gamma, x=e:E, \Gamma' \vdash b \rightarrow_\delta b'$ then $\Gamma, \Gamma'[x := e] \vdash b[x := e] \rightarrow_\delta b'[x := e]$.

Corollary 4.1.12.

1. If $\Gamma \vdash e \rightarrow_\delta e'$ and $\Gamma, x=e:E, \Gamma' \vdash B =_\delta B'$ then $\Gamma, x=e':E, \Gamma' \vdash B =_\delta B'$.

2. If $\Gamma, x=e:E, \Gamma' \vdash B =_\delta B'$ then $\Gamma, \Gamma'[x := e] \vdash B[x := e] =_\delta B'[x := e]$.
4.2 Church Rosser for $\beta, \delta$ and $\beta\delta$-reductions

Theorem 4.2.1 (Church Rosser for $\beta$-reduction). Let $\Gamma \in \mathcal{C}_3$ and $a, a_1, a_2 \in \mathcal{T}_3$ such that $\Gamma \vdash a \rightarrow_\beta a_1$ and $\Gamma \vdash a \rightarrow_\beta a_2$. Then there exists a term $a_3$ such that $\Gamma \vdash a_1 \rightarrow_\beta a_3$ and $\Gamma \vdash a_2 \rightarrow_\beta a_3$.

Proof: It follows from the fact that the Combinatorial Reduction System $<\mathcal{T}_3, \rightarrow_\beta>$ is orthogonal (See [Klo00]).

Definition 4.2.2. A mapping $| - | : \mathcal{T}_3 \times \mathcal{C}_3 \rightarrow \mathcal{T}$ is defined as follows.

$$
| x | \Gamma = \begin{cases} 
| a | \Gamma, & \text{if } \Gamma \equiv \langle \Gamma_1, x = a : \Gamma_2 \rangle \\
\varepsilon & \text{otherwise}
\end{cases}
$$

$$
| c | \Gamma = c \quad \text{for } c \in \mathcal{C}
$$

$$
| a \ b | \Gamma = | a | \Gamma \cdot | b | \Gamma
$$

$$
| \lambda x : A. \ a | \Gamma = \langle \lambda x : A | \Gamma. | a | \Gamma, x : A \rangle
$$

$$
| \Pi x : A. \ A | \Gamma = \langle \Pi x : A | \Gamma. | \Pi | \Gamma, x : A \rangle
$$

$$
| x = a : A \text{ in } b | \Gamma = | b | \Gamma [ x := | a | \Gamma] \quad \text{where } x \text{ is } \Gamma\text{-fresh}
$$

The value $| d | \Gamma$ is obtained from $d$ by unfolding all the definitions occurring in $\Gamma$ and in $d$. Note that $| d | \Gamma \in \mathcal{T}$ and $Def(\Gamma) \cap FV(| d | \Gamma) = \emptyset$. Conversely, if $d \in \mathcal{T}$ is such that $Def(\Gamma) \cap FV(d) = \emptyset$ then $| d | \Gamma = d$. Hence $| - |$ is a "projection" from $\mathcal{T}_3$ and $\mathcal{C}_3$ to $\mathcal{T}$. Later, it will be proved that $| d | \Gamma$ is the $\delta$-normal form of the term $d$.

Example 4.2.3. The value $| d | \Gamma$ is computed for one of the examples 3.2.3.

$$
| (id = (\lambda y : A. y) \cdot (\Pi y : A. A)) \cdot id | \Gamma = (\lambda x : (\Pi y : A. A) \cdot id) id
$$

The function $FV : \mathcal{T}_3 \rightarrow P(\mathcal{V})$ is extended to $\mathcal{C}_3 \times \mathcal{T}_3$.

Definition 4.2.4. The mapping $FV : \mathcal{C}_3 \times \mathcal{T}_3 \rightarrow P(\mathcal{V})$ is defined as follows:

$$
FV_{\varepsilon, A, \Gamma}(b) = FV(b)
$$

$$
FV_{\varepsilon, \Pi x, \Gamma}(b) = FV_{\varepsilon}(A) \cup (FV_{\varepsilon}(b) - \{x\})
$$

$$
FV_{\varepsilon, \lambda x, \Gamma}(b) = FV_{\varepsilon}(A) \cup FV_{\varepsilon}(b) \cup (FV_{\varepsilon}(b) - \{x\})
$$

Lemma 4.2.5.

1. If $x \notin FV(b)$ and $x$ is $\Gamma$-fresh then $x \notin FV(| b | \Gamma)$.

2. Let $< \Gamma_1, \Gamma_2, \Gamma_3 \in \mathcal{C}_3$ and $b \in \mathcal{T}_3$ be such that $(FV_{\varepsilon}(b)) \cap Def(\Gamma_2) = \emptyset$. Then

$$
| b | \Gamma_1, \Gamma_2, \Gamma_3 = | b | \Gamma_1, \Gamma_3
$$

3. Let $< \Gamma_1, x = a : A, \Gamma_2 \in \mathcal{C}_3$ and $a \in \mathcal{T}_3$. Then

$$
| a | \Gamma_1, x = a : A, \Gamma_2 = | a | \Gamma_1
$$

Proof:

1. It is proved by induction on the number of symbols occurring in $\Gamma$ and $b$.

2. It is proved by induction on the structure of $b$.

3. None of the variables in $Def(x = a : A, \Gamma_2)$ can occur in $a$. Hence the result follows immediately from the previous part.
Lemma 4.2.6. Let \( \Gamma, x := a : A \in \mathcal{C}_b \) and \( b \in T_b \). Then

\[
|b|_{\Gamma} [x := |a|_{\Gamma}] = |b[x := a]|_{\Gamma} = |b|_{\Gamma, x := a : A}
\]

Proof: It is proved by induction on the structure of \( b \). Only some cases are considered.

- Assume \( b \equiv x \).
  \[
  |x|_{\Gamma} [x := |a|_{\Gamma}] = x[x := |a|_{\Gamma}] = |a|_{\Gamma} \quad \text{(because \( x \) is \( \Gamma \)-fresh)}
  \]
  \[
  |x := a|_{\Gamma} = |a|_{\Gamma} \quad \text{(By definition of substitution)}
  \]
  \[
  \Gamma, x := a : A \quad \text{(By definition 4.2.2 part 1)}.
  \]

- Assume \( b \equiv y \neq x \).
  \[
  |y|_{\Gamma} [x := |a|_{\Gamma}] = |y|_{\Gamma} \quad \text{(By lemma 4.2.5)}
  \]
  \[
  |y := a|_{\Gamma} = |y|_{\Gamma} \quad \text{(By definition of substitution)}
  \]
  \[
  \Gamma, x := a : A = |y|_{\Gamma} \quad \text{(By lemma 4.2.5 part 2)}.
  \]

- Assume \( b \equiv (y = d : D \in e) \). Then \( y \) is \( \Gamma, x := a : A \)-fresh by convention 3.1.11.
  First the equality \( |b|_{\Gamma} [x := |a|_{\Gamma}] = |b[x := a]|_{\Gamma} \) will be proved.
  \[
  |b|_{\Gamma} [x := |a|_{\Gamma}] =
  \]
  \[
  |y = d : D \in e|_{\Gamma} [x := |a|_{\Gamma}] \quad \text{(By definition 4.2.2)}
  \]
  \[
  |y := |d|_{\Gamma} [x := |a|_{\Gamma}] = \quad \text{(By Lemma 4.1.7)}
  \]
  \[
  |y := |d|_{\Gamma} [x := |a|_{\Gamma}] \quad \text{(By definition of substitution)}
  \]

Next the equality \( |b|_{\Gamma} [x := |a|_{\Gamma}] = |b[x := a]|_{\Gamma} \) will be proved.

\[
|b[x := a]|_{\Gamma} = \quad \text{(By definition 4.2.2)}
\]

The following theorem states that a term reduces to its projection.

Lemma 4.2.7. Let \( \Gamma \in \mathcal{C}_b \) and \( d \in T_b \). Then

\[
\Gamma \vdash d \rightarrow_\delta |d|_{\Gamma}.
\]

Proof: It is proved by induction on the number of symbols occurring in \( \Gamma \) and in \( d \). Only two cases are considered.

- Assume \( d \equiv x \).
  If \( \Gamma \equiv \Gamma_1, x = b : B, \Gamma_2 \) then \( |x|_{\Gamma} = |b|_{\Gamma_1} \). By IH, it follows that \( \Gamma_1 \vdash b \rightarrow_\delta |b|_{\Gamma_1} \).

- Assume \( d \equiv (x := a : A \in b) \).
  By IH, it follows that \( \Gamma \vdash a \rightarrow_\delta |a|_{\Gamma}, \Gamma \vdash A \rightarrow_\delta |A|_{\Gamma} \) and that \( \Gamma \vdash b \rightarrow_\delta |b|_{\Gamma} \). By definition of \( \rightarrow_\delta \) and lemma 4.1.2, it follows that \( \Gamma \vdash (x := a : A \in b) \rightarrow_\delta (x := a : A \in b) \rightarrow_\delta |b|_{\Gamma} \).

By corollary 4.1.6, it follows that \(
\Gamma \vdash (x := a : A \in b) \rightarrow_\delta (x := a : A \in b) \rightarrow_\delta |b|_{\Gamma} \rightarrow_\delta |b|_{\Gamma} [x := |a|_{\Gamma}] \).

As \( |b|_{\Gamma} [x := |a|_{\Gamma}] = |x := a : A \in b|_{\Gamma} \), it follows that

\( \Gamma \vdash (x := a : A \in b) \rightarrow_\delta |x := a : A \in b|_{\Gamma} \).
The following theorem states that the projection of two terms that are in $\rightarrow_s$ are equal.

**Lemma 4.2.8.** Let $\Gamma \in C_0$ and $c, d \in T_0$ be such that $\Gamma \vdash c \rightarrow_s d$. Then

$$|c|_\Gamma = |d|_\Gamma.$$

**Proof:** It is proved by induction on the definition of $\rightarrow_s$. Only some cases are considered.

- Suppose that $\Gamma \vdash c \rightarrow_s d$ is $\Gamma, x : a : A \vdash x \rightarrow_s a$.

  By definition 4.2.2, it follows that $|x|_{\Gamma, x : a : A, \Gamma_2} = |a|_\Gamma$. By lemma 4.2.5, it follows that $|x|_{\Gamma_1, x : a : A, \Gamma_2} = |a|_\Gamma$. Then $|x|_{\Gamma_1, x : a : A, \Gamma_2} = |a|_{\Gamma_1, x : a : A, \Gamma_2}$.

- Suppose that $\Gamma \vdash c \rightarrow_s d$ is $\Gamma \vdash (x = a : A \in b) \rightarrow_s b$ if $x \notin FV(b)$.

  By definition 4.2.2, it follows that $|x = a : A|_\Gamma = |b|_\Gamma[x := |a|_\Gamma]$. By convention 3.1.11 the variable $x$ is $\Gamma$-fresh. Then by lemma 4.2.5 part 1 $|b|_\Gamma[x := |a|_\Gamma] = |b|_\Gamma$. Hence $|x = a : A \in b|_\Gamma = |b|_\Gamma$.

- Suppose that $\Gamma \vdash c \rightarrow_s d$ is $\Gamma \vdash (x = a : A \in b) \rightarrow_s (x = a' : A \in b')$ under the hypothesis $\Gamma \vdash a \rightarrow_s a'$.

  By definition 4.2.2, it follows that $|x = a : A|_\Gamma = |b|_\Gamma[x := |a|_\Gamma]$. By lemma 4.2.6, the following equalities hold $|b|_\Gamma[x := a : A] = |b|_\Gamma[x := |a|_\Gamma]$. By IH, the equality $|b|_\Gamma[x := a : A] = |b|_\Gamma[x := |a|_\Gamma]$ holds. Hence $|x = a : A \in b|_\Gamma = |x = a' : A \in b'|_\Gamma$.

- Suppose that $\Gamma \vdash c \rightarrow_s d$ is $\Gamma \vdash (x = a : A \in b) \rightarrow_s (x = a' : A \in b)$ under the hypothesis $\Gamma \vdash a \rightarrow_s a'$.

  By definition 4.2.2, it follows that $|x = a : A|_\Gamma = |b|_\Gamma[x := |a|_\Gamma]$. By IH, it follows that $|a|_\Gamma = |a'|_\Gamma$. Hence $|b|_\Gamma[x := |a|_\Gamma] = |b|_\Gamma[x := |a'|_\Gamma]$.

\[ \square \]

**Theorem 4.2.9 (Church Rosser for $\delta$-reduction).** Let $\Gamma \in C_0$ and $a, a_1, a_2 \in T_0$ such that $\Gamma \vdash a \rightarrow_s a_1$ and $\Gamma \vdash a \rightarrow_s a_2$. Then there exists a term $a_3$ such that $\Gamma \vdash a \rightarrow_s a_3$ and $\Gamma \vdash a_2 \rightarrow_s a_3$.

**Proof:** By theorem 4.2.7, $\Gamma \vdash a \rightarrow_s |a|_\Gamma$, $\Gamma \vdash a_1 \rightarrow_s |a_1|_\Gamma$ and $\Gamma \vdash a_2 \rightarrow_s |a_2|_\Gamma$.

By Lemma 4.2.8, it follows that $|a|_\Gamma = |a_1|_\Gamma = |a_2|_\Gamma$.

The scheme of the proof is illustrated in the following diagram.

\[ \square \]
The term $|a|_\Gamma$ is the 6-nf in $\Gamma$ of $a$, $a_1$ and $a_2$. This will be proved in corollary 4.3.3.

The following lemma states that the projection preserves $\beta$-reduction.

**Lemma 4.2.10.** Given $a, a' \in T_6$ and $\Gamma \in C_6$ such that $a \rightarrow_\beta a'$. Then

$$|a|_\Gamma \rightarrow_\beta |a'|_\Gamma.$$

**Proof:** It is proved by induction on the definition of $a \rightarrow_\beta a'$.

- **Assume $a \equiv (c \ d).$**
  There are 3 possibilities for the reduction $a \rightarrow_\beta a'$.
  
  1. Assume $a \rightarrow_\beta a'$ is $(\lambda x:B. \ b) \ d \rightarrow_\beta b[x := d]$. 
     By definition 4.2.2, it follows that
     $$|a|_\Gamma = (\lambda x:B. \ b)|_\Gamma = (\lambda x|_\Gamma[B]|_\Gamma. \ |b|_\Gamma)|_\Gamma.$$
     By Lemma 4.2.6, it follows that
     $$|b[x := d]|_\Gamma = |b|_\Gamma[x := |d|_\Gamma].$$
     Hence $|(\lambda x:B. \ b)|_\Gamma \rightarrow_\beta |b[x := d]|_\Gamma$.
     
  2. Assume $a \rightarrow_\beta a'$ is $(c \ d) \rightarrow_\beta (c' \ d')$ under the hypothesis $c \rightarrow_\beta c'$. 
     By IH, it follows that $|c|_\Gamma \rightarrow_\beta |c'|_\Gamma$. Hence $|a|_\Gamma \rightarrow_\beta |a'|_\Gamma$.
     
  3. Assume $a \rightarrow_\beta a'$ is $(c \ d) \rightarrow_\beta (c \ d')$.
     By IH, it follows that $|d|_\Gamma \rightarrow_\beta |d'|_\Gamma$. Hence $|c \ d|_\Gamma \rightarrow_\beta |c \ d'|_\Gamma$.

- **Assume $a \equiv (x=b:B \ in \ c).$**
  There are 3 possibilities for the reduction $a \rightarrow_\beta a'$.
  
  1. Assume $a \rightarrow_\beta a'$ is $(x=b:B \ in \ c) \rightarrow_\beta (x=b:B \ in \ c')$ under the hypothesis $c \rightarrow_\beta c'$. 
     By the definition 4.2.2, it follows that
     $$|a|_\Gamma = |x=b:B \ in \ c|_\Gamma = |c|_\Gamma[x := |b|_\Gamma].$$
     $$|a'|_\Gamma = |x=b:B \ in \ c'|_\Gamma = |c'|_\Gamma[x := |b|_\Gamma].$$
     By IH, it follows that $|c|_\Gamma \rightarrow_\beta |c'|_\Gamma$.
     By lemma 4.1.8, it follows that $|c|_\Gamma[x := |b|_\Gamma] \rightarrow_\beta |c'|_\Gamma[x := |b|_\Gamma]$.
     
  2. Assume $a \rightarrow_\beta a'$ is $(x=b:B \ in \ c) \rightarrow_\beta (x=b':B' \ in \ c)$ with $B \rightarrow_\beta B'$. 
     It follows that
     $$|a|_\Gamma = |c|_\Gamma[x := |b|_\Gamma] = |a'|_\Gamma.$$
     By lemma 4.2.10 and lemma 4.2.8, it follows that
     $$|c|_\Gamma[x := |b|_\Gamma] \rightarrow_\beta |c|_\Gamma[x := |b'|_\Gamma].$$
     
  3. Assume $a \rightarrow_\beta a'$ is $(x=b:B \ in \ c) \rightarrow_\beta (x=b':B \ in \ c)$ under the hypothesis $b \rightarrow_\beta b'$. 
     By the definition 4.2.2, the values of $|a|_\Gamma$ and $|a'|_\Gamma$ are
     $$|a|_\Gamma = |x=b:B \ in \ c|_\Gamma = |c|_\Gamma[x := |b|_\Gamma].$$
     $$|a'|_\Gamma = |x=b':B \ in \ c|_\Gamma = |c|_\Gamma[x := |b'|_\Gamma].$$
     By IH, it follows that $|b|_\Gamma \rightarrow_\beta |b'|_\Gamma$.
     By lemma 4.1.10, it follows that
     $$|c|_\Gamma[x := |b|_\Gamma] \rightarrow_\beta |c|_\Gamma[x := |b'|_\Gamma].$$

**Theorem 4.2.11 (Church Rosser for $\beta\delta$-reduction).** Let $\Gamma \in C_6$ and $a \in T_6$ be such that 

$$\Gamma \vdash a \rightarrow_\beta b \text{ and } \Gamma \vdash a \rightarrow_\delta c.$$ 

Then there exists a term $d \in T_6$ such that

$$\Gamma \vdash b \rightarrow_\delta d \text{ and } \Gamma \vdash c \rightarrow_\delta d.$$

**Proof:** By Lemma 4.2.7, it is deduced that

$$\Gamma \vdash a \rightarrow_\delta |a|_\Gamma, \ \Gamma \vdash b \rightarrow_\delta |b|_\Gamma \text{ and } \Gamma \vdash c \rightarrow_\delta |c|_\Gamma.$$

By lemma 4.2.10 and lemma 4.2.8, it follows that

$$\Gamma \vdash |a|_\Gamma \rightarrow_\beta |b|_\Gamma \text{ and that } \Gamma \vdash |a|_\Gamma \rightarrow_\beta |c|_\Gamma.$$

Hence the solid lines of the diagram are justified.
By Church Rosser theorem for \( \beta \)-reduction between pseudoterms in \( T \), the dotted arrows of the diagram are justified.

By composition along the borders, it follows that \( \Gamma \vdash b \rightarrow_{\beta \delta} d \) and \( \Gamma \vdash c \rightarrow_{\beta \delta} d \).

\[ \square \]

### 4.3 Weak and Strong Normalization for \( \rightarrow_{\delta} \)

**Definition 4.3.1.** Let \( a \in T_\delta \) and \( \Gamma \in C_\delta \). The term \( a \) is in \( \delta \)-normal form (or \( \delta \)-nf) in the context \( \Gamma \) if there is no term \( b \) such that \( \Gamma \vdash a \rightarrow_{\delta} b \).

A term can be in \( \delta \)-normal form but not in \( \beta \)-normal form, for example \((\lambda y:A. y)(\lambda y:A. y)\).

According to the following theorem, a term \( a \) is in \( \delta \)-nf in a context \( \Gamma \) if and only if \( a \) does not contain definitions and its free variables are not definitions in the context \( \Gamma \). As a corollary \( |a|_{\Gamma} \) is the \( \delta \)-nf of \( a \) in \( \Gamma \).

**Theorem 4.3.2.** Let \( a \in T_\delta \). Then

\[ a \text{ is in } \delta \text{-nf in } \Gamma \text{ if and only if } a \in T \text{ and } FV(a) \cap \text{Def}(\Gamma) = \emptyset. \]

**Proof:** It follows easily by induction on the structure of \( a \). \( \square \)

**Corollary 4.3.3. (Weak Normalization for \( \delta \)-reduction)** Let be \( \Gamma \in C_\delta \) and \( a \in T_\delta \), the term

\[ |a|_{\Gamma} \text{ is the } \delta \text{-nf of } a \text{ in } \Gamma. \]

**Proof:** By the previous theorem, it follows from \( |a|_{\Gamma} \in T \) and \( FV(|a|_{\Gamma}) \cap \text{Def}(\Gamma) = \emptyset \) that \( |a|_{\Gamma} \) is in \( \delta \)-nf in \( \Gamma \). By Church Rosser for \( \delta \)-reduction, the \( \delta \)-normal form is unique. \( \square \)

**Definition 4.3.4.** Let \( \Gamma \in C_\delta \) and \( b \in T_\delta \).

- The term \( b \) \( \delta \)-strongly normalizes in \( \Gamma \) if there is no infinite \( \delta \)-reduction starting with \( b \) in the context \( \Gamma \).

- The reduction \( \delta \) is strongly normalizing if for all pairs \( (\Gamma, b) \in C_\delta \times T_\delta \), the term \( b \) \( \delta \)-strongly normalizes in \( \Gamma \).

Note that by theorem 4.3.3 the normal form for an arbitrary \( \delta \)-term exists, but it is not guaranteed that all \( \delta \)-paths starting at the term are finite.

Next a function \( \text{nat}_{\text{me}}(-): C_\delta \times T_\delta \rightarrow \mathbb{N} \) that decreases with \( \delta \)-reduction is defined. It will be used to prove Strong Normalization for \( \rightarrow_{\delta} \).
Definition 4.3.5. The mapping \( \text{nat}(-) : C \times T \rightarrow \mathbb{N} \) is defined as follows.

\[
\begin{align*}
\text{nat}_{\Gamma,x = a:A, \Gamma_3}(x) &= \text{nat}_{\Gamma_1}(a) + 1 \\
\text{nat}_{\Gamma}(x) &= 0 \text{ if } x \notin \text{Def}(\Gamma) \\
\text{nat}_{\Gamma}(c) &= 0 \text{ if } c \in C \\
\text{nat}_{\Gamma}(x = a:A \text{ in } b) &= \text{nat}_{\Gamma}(a) + \text{nat}_{\Gamma}(A) + \text{nat}_{\Gamma,x = a:A}(b) + 1 \\
\text{nat}_{\Gamma}(a\ b) &= \text{nat}_{\Gamma}(a) + \text{nat}_{\Gamma}(b) \\
\text{nat}_{\Gamma}(\Pi x:A. \ a) &= \text{nat}_{\Gamma,x:A}(a) + \text{nat}_{\Gamma}(A) \\
\text{nat}_{\Gamma}(\lambda x:A. \ a) &= \text{nat}_{\Gamma,x:A}(a) + \text{nat}_{\Gamma}(A)
\end{align*}
\]

Lemma 4.3.6. Let \( \Gamma_1, \Gamma_2, \Gamma_3 \in C \) and \( b \in T \) be such that \( \text{FV}_{\Gamma_3}(b) \cap \text{Def}(\Gamma_2) = \emptyset \). Then

\[
\text{nat}_{\Gamma_1, \Gamma_2, \Gamma_3}(b) = \text{nat}_{\Gamma_1, \Gamma_3}(b).
\]

Proof: It is proved by induction on the number of symbols in \( \langle \Gamma_1, \Gamma_2, \Gamma_3 \rangle \) and in \( b \).

Lemma 4.3.7. Let \( \Gamma_1, x = a:A, \Gamma_2 \in C \) and \( b \in T \)

\[
\text{nat}_{\Gamma_1, x = a:A, \Gamma_2}(b) \geq \text{nat}_{\Gamma_1, \Gamma_3}(b).
\]

Proof: It is proved by induction on the number of symbols in \( \Gamma_1, x = a:A, \Gamma_2 \) and in \( b \). Only the case that \( b \) is a variable is considered.

- Assume \( b \equiv x \). Since \( \text{nat}_{\Gamma_1, \Gamma_3}(x) = 0 \) and \( \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(x) = \text{nat}_{\Gamma_1}(a) + 1 \), it follows that \( \text{nat}_{\Gamma_1, \Gamma_3}(x) < \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(x) \).

- Assume \( b \equiv y \) and \( y \in \text{Def}(\Gamma_1) \). Let \( \Gamma_1 \equiv \Gamma_1, y = c: C, \Gamma_1, 2 \). Since \( \text{nat}_{\Gamma_1, \Gamma_3}(y) = \text{nat}_{\Gamma_1}(c) + 1 \) and \( \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(y) = \text{nat}_{\Gamma_1}(c) + 1 \), it follows that \( \text{nat}_{\Gamma_1, \Gamma_3}(y) = \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(y) \).

- Assume \( b \equiv y \) and \( y \in \text{Def}(\Gamma_2) \). Let \( \Gamma_2 \equiv \Gamma_2, y = c: C, \Gamma_2, 2 \). Then \( \text{nat}_{\Gamma_1, \Gamma_3}(y) = \text{nat}_{\Gamma_1}(c) + 1 \) and \( \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(y) = \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(c) + 1 \).

- Assume \( b \equiv c \) and \( c \notin \text{Def}(\Gamma_1, x = a:A, \Gamma_2) \). Since \( \text{nat}_{\Gamma_1, \Gamma_3}(y) = 0 \) and \( \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(y) = 0 \), it follows that \( \text{nat}_{\Gamma_1, \Gamma_3}(y) = \text{nat}_{\Gamma_1, x = a:A, \Gamma_3}(y) \).

Definition 4.3.8. The concept "the context \( \Gamma \) \( \delta \)-reduces to the context \( \Gamma' \)" for \( \Gamma, \Gamma' \in C \) is written as \( \Gamma \vdash_{\delta} \Gamma' \) and is defined as follows.

\[
\begin{align*}
\Gamma \vdash E \rightarrow E' \\
\underline{\Gamma, y : E, \Gamma', \rightarrow_{\delta} \Gamma, y : E', \Gamma'} \\
\Gamma \vdash E \rightarrow E' \\
\underline{\Gamma, y = c : E, \Gamma', \rightarrow_{\delta} \Gamma, y = c : E', \Gamma'} \\
\Gamma \vdash c \rightarrow c' \\
\underline{\Gamma, y = c : E, \Gamma', \rightarrow_{\delta} \Gamma, y = c : E', \Gamma'}
\end{align*}
\]

Lemma 4.3.9. Let \( \Gamma \in C \) and terms \( d, d' \in T \) be such that \( \Gamma \vdash d \rightarrow_{\delta} d' \). Then

\[
\text{nat}_{\Gamma}(d) > \text{nat}_{\Gamma}(d').
\]

Proof: The following two properties are proved simultaneously by induction on number of symbols in \( \Gamma \) and in \( d \).
1. If $\Gamma \vdash d \rightarrow_\delta d'$ then $\text{natr}_\Gamma(d) > \text{natr}_\Gamma(d')$.

2. If $\Gamma \rightarrow_\delta \Gamma'$ then $\text{natr}_\Gamma(d) \geq \text{natr}_{\Gamma'}(d)$.

We only give the case that $d \equiv x$.

1. Then $\Gamma \vdash d \rightarrow_\delta d'$ is $\Gamma_1, x = a : A, \Gamma_2 \vdash x \rightarrow_\delta a$
   Since $\text{natr}_{\Gamma_1, x = a : A, \Gamma_2}(x) = \text{natr}_{\Gamma_1}(a) + 1$ and by lemma 4.3.6, $\text{natr}_{\Gamma_1}(a) = \text{natr}_{\Gamma}(a)$, it follows
   that $\text{natr}_{\Gamma}(x) > \text{natr}_{\Gamma}(a)$.

2. There are three possible cases. Suppose $\Gamma \equiv \Gamma_1, y = e : E, \Gamma_2$ with $\Gamma_1 \vdash e \rightarrow_\delta e'$ and $\Gamma_2 \equiv \Gamma_1, y = e' : E, \Gamma_2$.
   If $x \in \Gamma_1$ then $\text{natr}_{\Gamma}(x) = \text{natr}_{\Gamma}(x)$. If $x \equiv y$ then it follows from IH(1) that $\text{natr}_{\Gamma}(x) = \text{natr}_{\Gamma_1}(e) > \text{natr}_{\Gamma_1}(e') = \text{natr}_{\Gamma_1, y = e', E, \Gamma_2}(x)$. If $x \in \Gamma_2$ then it follows from IH(2) that $\text{natr}_{\Gamma}(x) \geq \text{natr}_{\Gamma}(x)$.
   If $x \not\in \Gamma$ then $\text{natr}_{\Gamma}(x) = x = \text{natr}_{\Gamma}(x)$.
   The rest of the cases are easy to prove.

\[ \Box \]

**Theorem 4.3.10. (Strong Normalization for $\delta$)** The reduction $\delta$ is strongly normalizing.

**Proof:** This follows directly from lemma 4.3.9.  

\[ \Box \]
Chapter 5
Properties of Well-Typed Terms

The properties in this chapter are proved for all well-typed terms, i.e. for terms \( a \) such that \( \exists A, \Gamma \vdash a : A \).

In Section 5.1, basic properties are proved for DPTS's. In Section 5.2 we prove that \( \lambda S \) is Weakly Normalizing if and only if \( \lambda S_b \) is Weakly Normalizing.

Finally Strong Normalization for \( \beta \delta \) reduction is proved for certain DPTS's.

5.1 Basic Properties

Lemma 5.1.1. (Generation Lemma)

1. \( \Gamma \vdash c : D \Rightarrow \exists s \in S[\Gamma \vdash D =_{\beta s} s \& c : s \in A] \)
2. \( \Gamma \vdash x : D \Rightarrow \exists s \in S \exists b \exists B =_{\beta s} D[\Gamma \vdash B : s \& x : B \in \Gamma \vdash b : B \& x = b : B \in \Gamma] \)
3. \( \Gamma \vdash (\Pi x : A. B) : D \Rightarrow \exists s_1, s_2, s_3 \in R[\Gamma \vdash A : s_1 \& \Gamma, x : A \vdash B : s_2 \& D =_{\beta s} s_3] \)
4. \( \Gamma \vdash (\lambda x : A. b) : D \Rightarrow \exists s \in S \exists B[\Gamma \vdash (\Pi x : A. B) : s \& \Gamma, x : A \vdash b : B \& D =_{\beta s} (\Pi x : A. B)] \)
5. \( \Gamma \vdash (b \ a) : D \Rightarrow \exists A, B[\Gamma \vdash b : (\Pi x : A. B) \& \Gamma \vdash a : A \& D =_{\beta s} B[x := a]] \)
6. \( \Gamma \vdash (x = a : A \text{ in } b) : D \Rightarrow \exists B[\Gamma, x = a : A \vdash b : B \& \Gamma \vdash (x = a : A \text{ in } B) : s \& D =_{\beta s} (x = a : A \text{ in } B) \vdash \Gamma, x = a : A \vdash b : s \& D =_{\beta s} s] \)

Proof: The different cases are all proved by induction on the derivation. 

Observe that in 6, the type of a term \( (x = a : A \text{ in } B) \) can be a sort \( s \) or an expression of the form \( (x = a : A \text{ in } B) \).

Lemma 5.1.2. (Correctness of types)

1. \( \Gamma \vdash A : B \Rightarrow \exists s \in S[B = s \lor \Gamma \vdash B : s] \)
2. \( \Gamma \vdash b : (\Pi x : A. B) \Rightarrow \exists s_1, s_2, s_3 \in R[\Gamma \vdash A : s_1 \& \Gamma, x : A \vdash B : s_2] \)

Proof:

1. It is proved by induction on the derivation of \( \Gamma \vdash A : B \).

2. It follows from the previous part that \( \Gamma \vdash (\Pi x : A. B) : s \). By the Generation Lemma part 3, it follows that \( \exists(s_1, s_2, s_3) \in R[\Gamma \vdash A : s_1 \& \Gamma, x : A \vdash B : s_2] \)

Lemma 5.1.3. If there exists \( b, B \in T_b \) such that \( \Gamma, x = a : A, \Gamma' \vdash b : B \) then \( \Gamma \vdash a : A \).
Proof: It is proved by induction on the derivation of $\Gamma, x=a: A, \Gamma' \vdash b : B$.

Lemma 5.1.4. (Thinning Lemma) Assume $\Gamma' \vdash b : B$. If $\Gamma \vdash a : A$ and $\Gamma \subseteq \Gamma'$ then $\Gamma \vdash a : A$

Proof: It is proved by induction on the derivation $\Gamma \vdash a : A$.

Lemma 5.1.5. If $\Gamma \vdash a : A$ and $\Gamma, x:A, \Gamma' \vdash b : B$ then $\Gamma, x=a:A, \Gamma' \vdash b : B$

Proof: It is proved by induction on the derivation of $\Gamma, x:A, \Gamma' \vdash b : B$.

Example 5.1.6. It is not true that $\Gamma, x=a:A, \Gamma' \vdash b : B \Rightarrow \Gamma, x=a:A, \Gamma' \vdash b : B$. For instance, we can derive $\alpha, x=\alpha, x, f, (\Pi y: \alpha. \alpha) \vdash (f y) : \alpha$.

But there is no term $B$ such that $\alpha, x=\alpha, x, f, (\Pi y: \alpha. \alpha) \vdash (f y) : B$.

Lemma 5.1.7. (Substitution Lemma)

1. If $\Gamma, x=a:A, \Gamma' \vdash b : B$ then $\Gamma, \Gamma'[x := a] \vdash b[x := a] : B[x := a]$

2. If $\Gamma \vdash a : A$ and $\Gamma, x:A, \Gamma' \vdash b : B$ then $\Gamma, \Gamma'[x := a] \vdash b[x := a] : B[x := a]$.

Proof:

1. It is proved by induction on the derivation of $\Gamma, x=a:A, \Gamma' \vdash b : B$.

2. It follows by lemma 5.1.5 and the previous part.

Definition 5.1.8. The concept "the context $\Gamma$ $\beta$-reduces to the context $\Gamma'$" is written as $\Gamma \rightarrow_{\beta} \Gamma'$ and is defined as follows.

\[
\begin{align*}
\Gamma \vdash E & \rightarrow_{\beta} E' \\
\Gamma, y:E, \Gamma' & \rightarrow_{\beta} \Gamma, y:E', \Gamma' \\
\Gamma \vdash E & \rightarrow_{\beta} E' \\
\Gamma, y=e:E, \Gamma' & \rightarrow_{\beta} \Gamma, y=e:E', \Gamma' \\
\Gamma \vdash e & \rightarrow_{\beta} e' \\
\Gamma, y=e:E, \Gamma' & \rightarrow_{\beta} \Gamma, y=e':E, \Gamma'
\end{align*}
\]

Theorem 5.1.9. (Subject Reduction Theorem) If $\Gamma \vdash d : D$ and $\Gamma \vdash d \rightarrow_{\beta} d'$ then $\Gamma \vdash d' : D$

Proof: The following properties are proved simultaneously by induction on the derivation of $\Gamma \vdash d : D$.

1. If $\Gamma \vdash d \rightarrow_{\beta} d'$ and $\Gamma \vdash d : D$ then $\Gamma \vdash d' : D$

2. If $\Gamma \rightarrow_{\beta} \Gamma'$ and $\Gamma \vdash d : D$ then $\Gamma \vdash d : D$

We only give the proof for some cases of the first property. Suppose that the last rule in the derivation of $\Gamma \vdash d : D$ is:

\[\begin{align*}
\Gamma \vdash b : (\Pi x:A. \ B) & \quad \Gamma \vdash a : A \\
\Gamma \vdash (b \ a) : B[x := a]
\end{align*}\]

There are 3 possibilities:

- Assume $\Gamma \vdash (b \ a) \rightarrow_{\beta} (b' \ a)$ under the hypothesis $\Gamma \vdash b \rightarrow_{\beta} b'$. By IH(1), it follows that $\Gamma \vdash b' : (\Pi x:A. \ B)$. By Application Rule, it follows that $\Gamma \vdash (b' \ a) : B[x := a]$.
\begin{itemize}
  \item Assume $\Gamma \vdash (b \ a) \rightarrow_{A'} (b \ a')$ under the hypothesis $\Gamma \vdash a \rightarrow_{A'} a'$.
  
  By IH(1), it follows that $\Gamma \vdash a' : A$. By Application Rule, it follows that $\Gamma \vdash (b \ a') : B[x := a']$. By lemma 4.1.10, it follows that $\Gamma \vdash B[x := a] \rightarrow_{A} B[x := a']$. By Correctness of Types, lemma 5.1.2 part 2, it follows that $\Gamma, x : A \vdash B : s_2$. By Substitution Lemma 5.1.7 part 2, it follows that $\Gamma \vdash B[x := a] : s_2$. By Conversion Rule, it follows that $\Gamma \vdash (b \ a') : B[x := a]$.

  \item Assume $\Gamma \vdash (\lambda x : A_1. b_1) a \rightarrow_{A} b_1[x := a]$. By Generation Lemma part 4, it follows that $A =_{A} A_1$ and that $\Gamma, x : A_1 \vdash b_1 : B$. By Generation Lemma part 3, it follows that $\Gamma \vdash A_1 : s$. By Conversion Rule, it follows that $\Gamma \vdash a : A_1$. By Substitution Lemma 5.1.7 part 2, it follows that $\Gamma \vdash b_1[x := a] : B[x := a]$.
\end{itemize}

\* ($\delta$ introduction) \[ \Gamma, x := a : A \vdash b : B \quad \Gamma \vdash (x := a : A) \in B : s \]

The following 5 cases are considered:

\begin{itemize}
  \item Assume $\Gamma \vdash (x := a : A) \in b \rightarrow_{A'} (x := a' : A')$ under the hypothesis $\Gamma \vdash A \rightarrow_{A'} A'$. By IH(2), it follows that $\Gamma, x := a : A' \vdash b : B$. By IH(1), it follows that $\Gamma \vdash (x := a' : A') \in b : B$. By the $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a' : A') \in b : (x := a' : A')$. By the Conversion Rule, it follows that $\Gamma \vdash (x := a' : A') \in b : (x := a' : A')$.

  \item Assume $\Gamma \vdash (x := a' : A) \in b \rightarrow_{A} (x := a' : A) \in b$ under the hypothesis $\Gamma \vdash a \rightarrow_{A} a'$. By IH(1), it follows that $\Gamma \vdash (x := a' : A) \in b : s$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a' : A) \in b : s$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a' : A) \in b : (x := a' : A)$. By Conversion Rule, it follows that $\Gamma \vdash (x := a' : A) \in b : (x := a' : A)$.

  \item Assume $\Gamma \vdash (x := a : A) \in b \rightarrow_{A} (x := a : A) \in b'$ under the hypothesis $\Gamma, x := a : A \vdash b' : B$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a : A) \in b' : (x := a : A)$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a : A) \in b' : (x := a : A)$.

  \item Assume $\Gamma \vdash (x := a : A) \in b \rightarrow_{A} (x := a : A) \in b$ under the hypothesis $\Gamma, x := a : A \vdash b : B$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a : A) \in b : (x := a : A)$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a : A) \in b : (x := a : A)$.

  \item Assume $\Gamma \vdash (x := a : A) \in b \rightarrow_{A} (x := a : A) \in b$ under the hypothesis $\Gamma, x := a : A \vdash b : B$. By $\delta$ introduction Rule, it follows that $\Gamma \vdash (x := a : A) \in b : (x := a : A)$.
\end{itemize}

\begin{definition}
\textbf{Definition 5.1.10.} The mapping $\mid \cdot \mid : C_5 \rightarrow C$ is defined as follows:
\begin{align*}
|e| & = e \\
|\Gamma, x : A| & = |\Gamma|, x : |A||r| \\
|\Gamma, x := a : A| & = |\Gamma|
\end{align*}

Note that if $\Gamma \in C$ then $|\Gamma| = \Gamma$. This mapping $\mid \cdot \mid$ is the projection from $C_5$ to $C$.
\end{definition}

\begin{theorem}
\textbf{Theorem 5.1.11.} Let $\Gamma \in C_4$ and $a, A \in T_4$.
\begin{align*}
\text{If } \Gamma \vdash_{\lambda S} a : A \text{ then } |\Gamma| \vdash_{\lambda S} |a||r| : |A||r|.
\end{align*}
\end{theorem}

\textbf{Proof:} Suppose that the last rule in the derivation of $\Gamma \vdash_{\lambda S} a : A$ is:

\begin{itemize}
  \item ($\delta$ start) $\vdash_{\lambda S} |\Gamma, x := a : A \vdash \delta_{\lambda S} x : A|$ where $x$ is $\Gamma$-fresh.
\end{itemize}

By IH $|\Gamma| \vdash_{\lambda S} |a||r| : |A||r|$. By definition 4.2.2 we have that $|x||r, x := a| = |a||r$. By lemma 4.2.5 part 2, it follows that $|A||r, x := a| = |A||r$. Hence $|\Gamma, x := a : A| \vdash_{\lambda S} |a||r, x := a| : |A||r, x := a|$. 

24
\(\Delta\). 
\begin{align*}
(\beta\text{-introduction}) 
\Gamma, x = a : A & \vdash_{\lambda S} b : B \\
\Gamma & \vdash_{\lambda S} (x = a : A \text { in } b) : (x = a : A \text { in } B)
\end{align*}

By IH \(\Gamma, x = a : A \vdash_{\lambda S} \|b\|_{x = a : A} : \|B\|_{x = a : A}\) and definition 4.2.2 we have that \(\|b\|_{x = a : A} = \|b\|_{x = a : \alpha} = \{x = a : A \text { in } |\!| b |\!|\} = x = a : A \text { in } |\!| b |\!|\) and that 
\(\|B\|_{x = a : A} = \|B\|_{x = a : \alpha} = \{x = a : A \text { in } |\!| B |\!|\} = x = a : A \text { in } |\!| B |\!|\).

Hence \(\Gamma \vdash_{\lambda S} (x = a : A \text { in } b) : (x = a : A \text { in } B)\).

\[\square\]

\textbf{Example 5.1.12.} In \((\lambda_\_)_S\) the following can be derived.

\[< \alpha : \ast, ID = (\Pi x : \alpha. \alpha) \ast, id = (\lambda x : \alpha. x) : ID > \vdash (\lambda_\_)_S \ast : id : ID\]

It follows from the previous theorem that

\[\alpha : \ast \vdash_{\lambda_\_} (\lambda x : \alpha. x) : (\Pi x : \alpha. \alpha)\]

All the definitions are unfolded and a derivation in \(\lambda_\_\) is obtained.

In the sense that there are not more typable terms with abbreviations, the inclusion of definitions to a PTS does not increase the strength of the system.

\textbf{Corollary 5.1.13.} (Conservativity) Let \(a \in T\) and \(\Gamma \in C\). Then

1. \(\exists A \ \Gamma \vdash_{\lambda S} a : A \iff \exists A \ \Gamma \vdash_{\lambda S_a} a : A\)
2. \(\exists A \ \Gamma \vdash_{\lambda S} a : a \iff \exists A \ \Gamma \vdash_{\lambda S_a} a : a\)
3. Let \(A \in T\). Then \(\Gamma \vdash_{\lambda S} a : A \iff \Gamma \vdash_{\lambda S_a} a : A\)

The first part of the previous corollary states that a term is typable in \(\lambda S\) iff it is typable in \(\lambda S_a\). The second part states that a type is inhabited in \(\lambda S\) iff it is inhabited in \(\lambda S_a\). If we interpret types as propositions and terms as proofs, this means that a proposition is provable in \(\lambda S\) iff it is provable in \(\lambda S_a\).

\textbf{Theorem 5.1.14 (Uniqueness of Types).} Let \(S\) be a singly sorted specification, \(\Gamma \in C\) and \(a, A, B \in T\) such that \(\Gamma \vdash_{\lambda S_a} a : A \text { and } \Gamma \vdash_{\lambda S_a} a : B\). Then \(\Gamma \vdash A \to B\).

\textbf{Proof:} It follows from lemma 5.1.11 that \(\Gamma \vdash_{\lambda S_a} |A|_{\Gamma} : |A|_{\Gamma}\) and \(\Gamma \vdash_{\lambda S_a} |A|_{\Gamma} : |B|_{\Gamma}\). As \(\lambda S\) is singly sorted then Uniqueness of Types holds. Hence \(|A|_{\Gamma} = |B|_{\Gamma}\). It follows from lemma 4.2.7 that \(\Gamma \vdash A \to \_ |A|_{\Gamma}\) and that \(\Gamma \vdash B \to \_ |B|_{\Gamma}\). Hence \(A =_{\beta\delta} B\).

\[\square\]

\textbf{5.2 Weak and Strong Normalization for \(\beta\delta\)-reduction}

In this section we will prove that if a PTS is weakly normalizing then the corresponding DPTS is weakly normalizing too.

Also we will prove that a DPTS is strongly normalizing if a slightly larger PTS is strongly normalizing. The idea of the proof of Strong Normalization is as follows:

- We define a mapping \(\{\_\}_\_ : T_\_ \times C_\_ \to T\) similar to the projection \(\_ |\_\). The value \(|a|_{\_}\) is a term that is obtained from \(a\) by unfolding all the definitions occurring in \(\_\) and in \(\_\). However \(\{\_\}_\_\) differs from \(\_ |\_\) in the value given to \(\{x=a : A \text { in } b\}_\_\). Instead of removing the local definition it is translated to a \(\beta\)\(\delta\)-redex, i.e. an application and an abstraction.

- This function \(\{\_\}_\_\) maps an infinite \(\beta\delta\) reduction sequence to an infinite \(\beta\) reduction sequence.

- The function \(\{\_\}_\_\) maps terms that are typable in a DPTS \(\lambda S_\_\) to terms that are typable in a PTS slightly larger than \(\lambda S\).
Definition 5.2.1. Let $\Gamma \in C_\delta$ and $b \in T_\delta$. The term $b$ $\beta\delta$-weakly normalizes in $\Gamma$ if there is a term $d$ in $\beta\delta$-normal form in $\Gamma$ such that $\Gamma \vdash b \rightarrow_{\beta\delta} d$.

Definition 5.2.2. Let $\lambda S_\delta$ be a DPTS and $\rightarrow_{\rho}$ a reduction relation. Then $\lambda S_\delta$ is $\rho$-weakly normalizing if $\rho$-weakly normalize in $\Gamma$ for all $a \in T_\delta$ and $\Gamma \in C_\delta$ such that $\Gamma \vdash_{\lambda S_\delta} a : A$.

Theorem 5.2.3. Let $\lambda S$ be a PTS. $\lambda S$ is $\beta$-weakly normalizing if and only if $\lambda S_\delta$ is $\beta\delta$-weakly normalizing.

Proof: Suppose $\Gamma \vdash_{\lambda S_\delta} a : A$. Then $|\Gamma|_{\lambda S} \vdash |a|_{\Gamma} : |A|_{\Gamma}$. There exists $b \in T$ such that $|a|_{\Gamma} \rightarrow_{\beta} b$. Hence $\Gamma \vdash a \rightarrow_{\rho} |a|_{\Gamma} \rightarrow_{\beta} b$.

The converse is obvious. \qed

Definition 5.2.4. Let $\Gamma \in C_\delta$ and $b \in T_\delta$. The term $b$ $\beta\delta$-strongly normalizes in $\Gamma$ if there is no infinite $\beta\delta$-reduction starting at $b$ in $\Gamma$.

Definition 5.2.5. Let $\lambda S_\delta$ be a DPTS and $\rightarrow_{\rho}$ a reduction relation. Then $\lambda S_\delta$ is $\rho$-strongly normalizing if $\rho$-strongly normalize in $\Gamma$ for all $a \in T_\delta$ and $\Gamma \in C_\delta$ such that $\Gamma \vdash_{\lambda S_\delta} a : A$.

The function $| \_ |_{\_}$ may not map infinite $\beta$ reduction sequences to infinite $\beta$ reduction sequences as the following example shows:

Example 5.2.6. Suppose there is an infinite $\beta$-reduction sequence starting at $a$ and hence at $(x = a : A$ in $b)$. If $x \notin FV(b)$ and $b$ is a $\beta$ normal form then there is no $\beta$-reduction sequence starting at $| x = a : A$ in $b |_{\equiv}$.

This means that the function $| \_ |_{\_}$ cannot be used to establish a relation between PTS's and DPTS's that are strongly normalizing. A new function $| \_ |_{\_}$ will be defined.

Definition 5.2.7. The mapping $| \_ |_{\_} : T_\delta \times C_\delta \rightarrow T$ is defined as follows:

$$
\begin{aligned}
{x}_{\Gamma} &= \begin{cases} 
{a}_{\Gamma}, & \text{if } \Gamma \equiv \Gamma_1, x = a : A, \Gamma_2 \\
\Gamma, & \text{otherwise} 
\end{cases} \\
{c}_{\Gamma} &= c, & \text{if } c \in C \\
{a b}_{\Gamma} &= {a}_{\Gamma}{b}_{\Gamma} \\
{x a : A}_{\Gamma} &= {x a : A}_{\Gamma} \\
{\Pi x : A}_{\Gamma} &= {\Pi x : A}_{\Gamma} \\
{x = a : A}_{\Gamma} &= (\lambda x : A)_{\Gamma} \{ b \}_{\Gamma,x : A} \\
{\lambda x : A}_{\Gamma} &= ( \lambda x : A)_{\Gamma} \{ b \}_{\Gamma,x : A} \\
{\lambda a : A}_{\Gamma} &= ( \lambda a : A)_{\Gamma} \{ b \}_{\Gamma,x = a : A}(a)_{\Gamma}
\end{aligned}
$$

Like $| \_ |_{\_}$, the value $| a |_{\Gamma}$ is the unfolding of all the definitions occurring in $\Gamma$ in $a$. However $| \_ |_{\_}$ differs from $| \_ |_{\_}$ in the value given for $(x = a : A$ in $b)$. Instead of removing the local definition it is translated to a $\beta$-redex, an application and an abstraction.

Example 5.2.8. Recall that in example 1.3 we show that $(\lambda x : A. b)a$ may not be typable when $(x = a : A$ in $b)$ is. Let $c \equiv \lambda a : * \cdot (x = a : * \cdot \lambda y : x. \lambda f : \alpha \rightarrow \alpha. fy) \alpha$ be the term used in example 1.3. The corresponding term expressed as an application and an abstraction is not typable in any system of the $\lambda$-cube. But

$$
\begin{aligned}
{c}_{\Gamma} &\equiv \lambda a : * \cdot (\lambda x : * \cdot \lambda y : \alpha. \lambda f : \alpha \rightarrow \alpha. fy) \alpha \\
\end{aligned}
$$

is typable in $\lambda \Sigma$. This is because the definition of $x$ is unfolded by $a$ and then $x$ does not occur in the expression $\lambda y : \alpha. \lambda f : \alpha \rightarrow \alpha. fy$. That is because $\lambda x : * \cdot \lambda y : \alpha. \lambda f : \alpha \rightarrow \alpha. fy$ is typable in $\lambda \Sigma$. This is because the definition of $x$ is unfolded by $a$ and then $x$ does not occur in the expression $\lambda y : \alpha. \lambda f : \alpha \rightarrow \alpha. fy$.

The mapping $| \_ |_{\_}$ is extended to contexts.
Definition 5.2.9. The mapping \(\{-\} : C \rightarrow C\) is defined as follows:

\[
\begin{align*}
\{c\} &= c \\
\{\Gamma, x : A\} &= \{\Gamma\}, x : \{A\}r \\
\{\Gamma, x = a : A\} &= \{\Gamma\}, x : \{A\}r
\end{align*}
\]

Similar properties proved for the projection \(\lfloor . \rfloor\) will be proved for the function \(\{-\}\).

Lemma 5.2.10.

1. If \(x\) is \(\Gamma\)-fresh and \(x \notin FV(b)\) then \(x \notin FV(\{b\}r)\).
2. Let \(<\Gamma_1, \Gamma_2, \Gamma_3 \in C_3\) and \(b \in T_b\) be such that \((FV_b(b)) \cap Def(\Gamma_2) = \emptyset\). Then

\[
\{b\}_{\Gamma_1, \Gamma_2, \Gamma_3} \equiv \{b\}_{\Gamma_1, \Gamma_3}.
\]

3. Let \(<\Gamma_1, y = a : A, \Gamma_2 \in C_3\). Then \(\{a\}_{\Gamma_1, y = a : A, \Gamma_2} \equiv \{a\}_{\Gamma_1}\).

Proof:

It follows from the previous part.

Lemma 5.2.11. \(\{b\}_{r[x := a]} \equiv \{b[x := a]\}r \equiv \{b\}_{r, x = a : A}\)

Proof: Induction on the structure of \(b\).

By the following lemma, \(\{-\}\) maps an infinite \(\beta\delta\) reduction sequence to an infinite \(\beta\) reduction sequence.

Lemma 5.2.12.

1. If \(c \rightarrow_{\beta} d\) then \(\{c\}_{\Gamma} \rightarrow_{\beta} \{d\}_{\Gamma}\).
2. If \(\Gamma \vdash c \rightarrow_{\delta} d\) then \(\{c\}_{\Gamma} \rightarrow_{\beta} \{d\}_{\Gamma}\).

Proof: Both parts are proved by induction on the structure of \(c\).

1. Suppose \(c\) is an application, \(c \equiv (a \ b)\). There are two possibilities for the reduction \(c \rightarrow_{\beta} d\):

   - \(d \equiv (a' \ b')\) with \((a \rightarrow_{\beta} a' \land b \equiv b')\) or \((a \equiv a' \land b \rightarrow_{\delta} b')\). Then it follows immediately from the IH that \(\{a\}r\{b\}r \rightarrow_{\beta} \{a'\}r\{b'\}r\).
   - \(a \equiv (\lambda x : A. \ B)\) and \(d \equiv B[x := b]\). Then

\[
\begin{align*}
\{c\}_{\Gamma} &\equiv \{a\}_{\Gamma}\{b\}_{\Gamma} \\
&\equiv (\lambda x : \{A\}_{\Gamma} \{B\}_{\Gamma})\{b\}_{\Gamma} \\
&\rightarrow_{\beta} \{B\}_{\Gamma}[x := \{b\}_{\Gamma}] \\
&\equiv \{B[x := b]\}_{\Gamma} \quad \text{by lemma 5.2.11} \\
&\equiv \{d\}_{\Gamma}
\end{align*}
\]

2. Only two cases are considered.

   (a) Suppose \(c \equiv (x = a : A\ \text{in} \ b)\). There are 4 possibilities for the reduction \(c \rightarrow_{\delta} d\):

   - \(d \equiv (x = a : A' \ \text{in} \ b)\) with \(\Gamma \vdash A \rightarrow_{\delta} A'\). Then it follows from the IH that

\[
\begin{align*}
(\lambda x : \{A\}_{\Gamma} \{b\}_{\Gamma, x = a : A})\{a\}_{\Gamma} \rightarrow_{\beta} (\lambda x : \{A'\}_{\Gamma} \{b\}_{\Gamma, x = a : A})\{a\}_{\Gamma}.
\end{align*}
\]
Suppose $d \equiv (x = a : A \in b)$ with $\Gamma \vdash a \rightarrow e a'$. Then it follows from the IH and lemma 4.1.10 that
\[
\begin{align*}
(\lambda x : A) r \cdot \{b \mid x := \{a\} r\} r &= \frac{\beta}{(\lambda x : A) r \cdot \{b \mid x := \{a'\} r\} r}
\end{align*}
\]

Suppose $x \equiv (x = a : A \in b')$ with $\Gamma, z = a : A \vdash b \rightarrow e b'$ Then it follows from the IH that
\[
\begin{align*}
(\lambda x : A) r \cdot \{b \mid x := \{a\} r\} r &= \frac{\beta}{(\lambda x : A) r \cdot \{b' \mid x := \{a'\} r\} r}
\end{align*}
\]

Suppose $x \notin FV(b)$ and $d \equiv b$. Then it follows from the IH that
\[
\begin{align*}
(\lambda x : A) r \cdot \{b \mid x := \{a\} r\} r &= \frac{\beta}{(\lambda x : A) r \cdot \{b' \mid x := \{a'\} r\} r}
\end{align*}
\]

(b) Suppose $c \equiv x$. This means that $\Gamma \equiv \Gamma_1, x = d : D, \Gamma_2$ and $\Gamma \vdash x \rightarrow e d$.

By the definition of $\{\_\}$, and lemma 5.2.10 part 3 we have that $\{x\} r_1, x = d : D, r_2 \equiv \{d\} r_1$, and $\{d\} r_1, x = d : D, r_2$.

\[\square\]

**Corollary 5.2.13.**

1. If $\Gamma \vdash B \rightarrow e B'$, then $\{B\} r \rightarrow \beta \{B'\} r$.
2. If $\Gamma \vdash B =_{\beta} B'$, then $\{B\} r =_{\beta} \{B'\} r$.

**Proof:** These follow from part 2 of the previous lemma. \[\square\]

**Definition 5.2.14.** The specification $S = (S, A, R)$ is called quasi-full if for all $s_1, s_2 \in S$ there exists $s_3 \in S$ such that $(s_1, s_2, s_3) \in R$. \[1\]

**Definition 5.2.15.** Let $S = (S, A, R)$ and $S' = (S', A', R')$ be such that

1. $S \subseteq S'$, $A \subseteq A'$, and $R \subseteq R'$
2. $S'$ is quasi-full
3. for all $s \in S$ there is a sort $s' \in S'$ such that $(s : s') \in A'$ (i.e. the sorts of $S$ are not topsorts in $S'$).

Then the specification $S'$ is called a completion of $S$.

**Example 5.2.16.** The system $\lambda C_{\infty}$ is a completion of $\lambda C$, $\lambda HOL$ and itself.

This definition is necessary in order to prove that $\{\_\}_-$ maps terms that are typable in a DPTS $\lambda S_k$ to terms that are typable in a slightly larger PTS, i.e. a PTS $\lambda S'$ with $S'$ a completion of $S$. Remember that $\{\_\}_-$ translates a local definition to a $\lambda$-abstraction with an argument:
\[
\{x = a : A \mid b\} r = (\lambda x : A) r \cdot \{b \mid x := \{a\} r\} r
\]

Condition 2 is necessary to ensure that all these $\lambda$-abstractions introduced by $\{\_\}_-$ are allowed in $S'$. The typing of the abstraction is restricted by the set $R$ of rules whereas the typing of $(x = a : A \in b)$ is not. Let $e \equiv (x = a : x \in \lambda y : x. \lambda f : x \rightarrow x. \lambda g) b$ be the term in the example 1.4. This term is typable in the system $(\lambda _-, _-) \_b$ but the term $\{e\}, \equiv (\lambda x : a. \lambda y : a. \lambda f : a \rightarrow a. \lambda g) a$ is not typable in $\lambda _-$. 

\[1\]Note that if a specification is full then it is quasi-full. But the converse is not true. A specification is full if for all $s_1, s_2 \in S (s_1, s_2, s_2) \in R$.
Condition 3 is necessary because we can not type these abstractions introduced by \( \{ \_ \} \text{ } \) if \( A \) is a toposort. For example the term \( (x = *: \Box \text{ in } x) \) is typable in \( \lambda C_S \) but
\[
\{ x = *: \Box \text{ in } x \}_r \equiv (\lambda x: \Box. \ast)_r
\]
is not typable in \( \lambda C \).

**Lemma 5.2.17.** Let \( c \in C, a, A \in T_6 \) and \( \Gamma \in T_6 \). If \( c \) occurs in \( A \) or \( a \) or \( \Gamma \) and \( \Gamma \vdash_{\lambda S_e} a : A \) then \( c \in S \).

**Proof:** By induction on the derivation of \( \Gamma \vdash_{\lambda S_e} a : A \).

The next lemma states that any \( \lambda S' \) type that is in the range of \( \{ \_ \}_r \) cannot be one of the toposorts of \( \lambda S' \):

**Lemma 5.2.18.** Let \( S = (S, A, R) \) and \( S' = (S', A', R') \) be such that \( S' \) is a completion of \( S \). Then if \( \Delta \vdash_{\lambda S_e} a : A \) and \( \Delta \vdash_{\lambda S'} \{ a \}_r : \{ A \}_r \) then \( \Delta \vdash_{\lambda S'} \{ A \}_r : s \).

**Proof:** Assume \( \Delta \vdash_{\lambda S'} \{ a \}_r : \{ A \}_r \). By correctness of types \( \Delta \vdash_{\lambda S_e} \{ A \}_r : s' \) or \( \{ A \}_r \equiv s \). Suppose \( \{ A \}_r \equiv s \). Since \( \Delta \vdash_{\lambda S_e} a : A \) and by lemma 5.2.17 we have that \( s \in S \). By (3) there is a sort \( s' \in S' \) such that \( (s : s') \in A' \) and hence \( \Delta \vdash_{\lambda S'} s : s' \), i.e. \( \Delta \vdash_{\lambda S'} \{ A \}_r : s' \).

The next theorem states that \( \{ \_ \}_r \) maps terms that are typable in a DPTS \( \lambda S_e \) to terms that are typable in a slightly larger PTS, i.e. \( \lambda S' \) with \( S' \) a completion of \( S \).

**Theorem 5.2.19.** Let \( S = (S, A, R) \) and \( S' = (S', A', R') \) be such that \( S' \) is a completion of \( S \). Then \( \Gamma \vdash_{\lambda S_e} a : A \Rightarrow \{ \Gamma \} \vdash_{\lambda S'} \{ a \}_r : \{ A \}_r \).

**Proof:** By induction on the derivation of \( \Gamma \vdash_{\lambda S_e} a : A \). Suppose the last step in the derivation is

- (6-start) \( \frac{\Gamma \vdash_{\lambda S_e} a : A}{\Gamma, x = a : A \vdash_{\lambda S_e} A} \)
  
  By IH \( \{ \Gamma \} \vdash_{\lambda S'} \{ a \}_r : \{ A \}_r \). By definition we have that \( \{ x \}_r, x = a, A = \{ a \}_r \). By lemma 5.2.10 part 2 we have that \( \{ A \}_r, x = a, A = \{ A \}_r \). It follows from lemma 5.2.18 that \( \{ \Gamma \} \vdash_{\lambda S'} \{ a \}_r : s \).

- (6-weakening) \( \frac{\Gamma \vdash_{\lambda S_e} b : B}{\Gamma, x = a : A \vdash_{\lambda S_e} b : B} \)
  
  By IH \( \{ \Gamma \} \vdash_{\lambda S'} \{ b \}_r : \{ B \}_r \) and \( \{ \Gamma \} \vdash_{\lambda S'} \{ a \}_r : \{ A \}_r \). Then by lemma 5.2.18 \( \{ \Gamma \} \vdash_{\lambda S'} \{ A \}_r : s \).

  By weakening rule \( \{ \Gamma \}, x : \{ A \}_r \vdash_{\lambda S'} \{ b \}_r : \{ B \}_r \).

  We have that \( x \not\in \text{FV}(b) \) and \( x \not\in \text{FV}(B) \). By lemma 5.2.10 \( \{ b \}_r, x = a, A = \{ b \}_r \) and \( \{ B \}_r, x = a, A = \{ B \}_r \).

- (6-conversion) \( \frac{\Gamma \vdash_{\lambda S_e} b : B}{\Gamma \vdash_{\lambda S_e} B' : s} \)
  
  It follows from IH that \( \{ \Gamma \} \vdash_{\lambda S'} \{ b \}_r : \{ B \}_r \) and \( \{ \Gamma \} \vdash_{\lambda S'} \{ B' \}_r : s \). By corollary 5.2.13 part 2 it follows from \( \Gamma \vdash B =_s B' \) that \( \{ B \}_r =_\beta \{ B' \}_r \). Then using \( \beta \)-conversion rule \( \{ \Gamma \} \vdash_{\lambda S'} \{ b \}_r : \{ B' \}_r \).

- (6-formation) \( \frac{\Gamma, x = a : A \vdash_{\lambda S_e} B : s}{\Gamma \vdash_{\lambda S_e} \{ x = a : A \} \vdash_{\lambda S' e} B : s} \)
  
  By IH

\[
\{ \Gamma, x = a : A \} \vdash_{\lambda S'} \{ B \}_r, x = a, A = \{ B \}_r \quad \text{(5.1)}
\]
The derivation of $\Gamma, x = a : A \vdash_{\lambda S_b} B : s$ contains a (shorter) derivation of $\Gamma \vdash_{\lambda S_b} a : A$, so also by IH

$$\{\Gamma\} \vdash_{\lambda S'} \{a\}_r : \{A\}_r$$ \hspace{1cm} (5.2)

By lemma 5.2.18 it follows from (5.1) and (5.2) that there are $s_1, s_2 \in S'$ such that

$$\{\Gamma, x = a : A\} \vdash_{\lambda S'} x : s_2$$

$$\{\Gamma\} \vdash_{\lambda S'} \{A\}_r : s_1$$ \hspace{1cm} (5.4)

The following is a derivation of $\{\Gamma\} \vdash_{\lambda S'} \{x = a : A \in B\}_r : s$.

$$\{\Gamma\} \vdash_{\lambda S'} (\Pi x \in \{A\}_r. s) : s_3$$ \hspace{1cm} (prod)

$$\{\Gamma\} \vdash_{\lambda S'} (\lambda x \in \{A\}_r. (B)_{r,x=e:A}) : (\Pi x \in \{A\}_r. s)$$ \hspace{1cm} (abs)

$$\{\Gamma\} \vdash_{\lambda S'} (\lambda x \in \{A\}_r. (B)_{r,x=e:A})(a)_r : s \equiv a \{a\}_r$$ \hspace{1cm} (app)

and $(\lambda x \in \{A\}_r. (B)_{r,x=e:A})(a)_r \equiv \{x = a : A \in B\}_r$.

* $\delta$-introduction

$$\Gamma, x = a : A \vdash_{\lambda S_b} b : B \Gamma \vdash_{\lambda S_b} (x = a : A \in B)_r : s$$

$$\Gamma \vdash_{\lambda S_b} (x = a : A \in B)_r : s \equiv \{x = a : A \in B\}_r$$

By IH

$$\{\Gamma, x = a : A\} \vdash_{\lambda S'} (b)_{r,x=e:A} : (B)_{r,x=e:A}$$

$$\{\Gamma\} \vdash_{\lambda S'} \{x = a : A \in B\}_r : s$$ \hspace{1cm} (5.2)

The derivation of $\Gamma, x = a : A \vdash_{\lambda S_b} b : B$ contains a (shorter) derivation of $\Gamma \vdash_{\lambda S_b} a : A$, so also by the IH

$$\{\Gamma\} \vdash_{\lambda S'} \{a\}_r : \{A\}_r$$ \hspace{1cm} (5.3)

By lemma 5.2.18 it follows from (5.1) and (5.3) that there are $s_1, s_2 \in S'$ such that

$$\{\Gamma, x = a : A\} \vdash_{\lambda S'} \{B\}_{r,x=e:A} : s_2$$

$$\{\Gamma\} \vdash_{\lambda S'} \{A\}_r : s_1$$ \hspace{1cm} (5.5)

Then

$$\{\Gamma\} \vdash_{\lambda S'} (\Pi x \in \{A\}_r. (B)_{r,x=e:A}) : s_3$$ \hspace{1cm} (prod)

$$\{\Gamma\} \vdash_{\lambda S'} (\lambda x \in \{A\}_r. (B)_{r,x=e:A})(a)_r : s \equiv a \{a\}_r$$ \hspace{1cm} (app)

so using the conversion rule

$$\{\Gamma\} \vdash_{\lambda S'} (x = a : A \in B)_r : \{B\}_{r,x=e:A}(a)_r \equiv \{x = a : A \in B\}_r$$ \hspace{1cm} (5.6)

The rest of the cases are easy to prove. \hspace{1cm} $\square$

**Theorem 5.2.20.** Let $S = (S, A, R)$ and $S' = (S', A', R')$ be such that $S'$ is a completion of $S$. If the PTS $\lambda S'$ is $\beta$-strongly normalizing, then the DPTS $\lambda S_b$ is $\beta\delta$-strongly normalizing.
Proof: Suppose that \( \lambda S' \) is \( \beta \)-strongly normalizing, and suppose towards a contradiction that \( \lambda S_4 \) is not \( \beta \delta \)-strongly normalizing, i.e. there is an infinite \( \beta \delta \)-reduction sequence starting at \( a \) and \( \Gamma \vdash_{\lambda S_4} a : A \).

Observe that the number of \( \beta \)-reductions in this sequence is infinite, i.e. \( \forall n \in \mathbb{N} \ \exists m > n : \Gamma \vdash a_m \rightarrow_{\beta} a_{m+1} \). Otherwise it would follow that there is \( n_0 \in \mathbb{N} \) such that \( \forall m > n_0 \ \Gamma \vdash a_m \rightarrow_{\beta} a_{m+1} \). Hence the sequence \( a_{n_0+1} \rightarrow_{\beta} a_{n_0+2} \rightarrow_{\beta} \ldots \) would be infinite. As \( \delta \) is strongly normalizing, this can not happen. Hence the number of \( \beta \)-reduction steps in the sequence \( \Gamma \vdash a \rightarrow_{\beta \delta} a_1 \rightarrow_{\beta \delta} a_2 \rightarrow_{\beta \delta} \ldots \) is infinite. Then this sequence is of the form

\[
\Gamma \vdash a \rightarrow_{\beta} a_{n_1} \rightarrow_{\beta} a_{n_2} \rightarrow_{\beta} a_{n_3} \rightarrow_{\beta} \ldots
\]

By lemma 5.2.12 and corollary 5.2.13 part 1 there is an infinite \( \beta \)-reduction sequence starting at \( \{a\}_\Gamma \):

\[
\{a\}_\Gamma \rightarrow_{\beta} \{a_{n_1}\}_\Gamma \rightarrow_{\beta} \{a_{n_2}\}_\Gamma \rightarrow_{\beta} \{a_{n_3}\}_\Gamma \rightarrow_{\beta} \ldots
\]

and by lemma 5.2.19 \( \{\Gamma\} \vdash_{\lambda S'} \{a\}_\Gamma : \{A\}_\Gamma \), which contradicts the assumption that \( \lambda S' \) is \( \beta \)-strongly normalizing.

Corollary 5.2.21. The following systems are strongly normalizing:

1. The system \( \lambda C_\infty \) extended with definitions, i.e. \( (\lambda C_\infty)_I \).
2. The calculus of constructions extended with definitions, i.e. \( \lambda C_\delta \).
3. The system of higher order logic extended with definitions, i.e. \( \lambda HOL_\delta \) is strongly normalizing.

Proof: The system ECC (see [Luo]) is strongly normalizing and contains \( \lambda C_\infty \). Then \( \lambda C_\infty \) is strongly normalizing too. This system \( \lambda C_\infty \) is a completion of itself. Hence it follows from the previous theorem that \( \lambda C_\infty \) extended with definitions is strongly normalizing. Since \( (\lambda C_\infty)_I \) contains \( \lambda C_\delta \) and \( \lambda HOL_\delta \), part 2 and 3 follow from part 1.

Theorem 5.2.20 is somewhat unsatisfactory. It would be nicer to prove a stronger property, namely that a DPTS \( \lambda S_4 \) is \( \beta \delta \)-strongly normalizing if the PTS \( \lambda S \) is \( \beta \)-strongly normalizing. On the other hand, we do not know any strongly normalizing PTS \( \lambda S \) for which theorem 5.2.20 cannot be used to prove strong normalization of \( \lambda S_4 \). In particular, all strongly normalizing PTS’s given in [Bar92] have a completion that is \( \lambda C_\infty \).

31
Appendix A

Pure Type Systems with definitions

Definition A.1.
The set $T_0$ of pseudoterms and the set $C_0$ of contexts is given by

$$T_0 ::= V \mid C \mid (T_0, T_0) \mid (V : T_0, T_0) \mid (V = T_0 : T_0 \text{ in } T_0)$$

$$C_0 ::= \varepsilon \mid < C_0, V : T_0 > \mid < C_0, V = T_0 : T_0 >$$

where $V$ is the set of variables and $C$ is the set of constants.

Definition A.2.
Let be $\Gamma \in C_0$, and $a, a' \in T_0$. The concept "the term $a$ $\delta$-reduces to the term $a'$ in the context $\Gamma$" is written as $\Gamma \vdash a \rightarrow_\delta a'$ and is defined by the following rules:

$$\Gamma_1, x = a : A, \Gamma_2 \vdash x \rightarrow_\delta a$$

$$\Gamma \vdash (x = a : A \text{ in } b) \rightarrow_\delta b$$

if $x \notin \text{FV}(b)$

$$\Gamma, x = a : A \vdash b \rightarrow_\delta b'$$

$$\Gamma \vdash (x = a : A \text{ in } b) \rightarrow_\delta (x = a : A \text{ in } b')$$

$$\Gamma \vdash a \rightarrow_\delta a'$$

$$\Gamma \vdash (x = a : A \text{ in } b) \rightarrow_\delta (x = a' : A \text{ in } b)$$

$$\Gamma \vdash b \rightarrow_\delta b'$$

$$\Gamma \vdash (ab) \rightarrow_\delta (a'b)$$

$$\Gamma, x : A \vdash a \rightarrow_\delta a'$$

$$\Gamma, x : A \vdash (ab) \rightarrow_\delta (a'b)$$

$$\Gamma \vdash (\lambda x : A. a) \rightarrow_\delta (\lambda x : A'. a')$$

$$\Gamma \vdash a \rightarrow_\delta a'$$

$$\Gamma \vdash (\Pi x : A. a) \rightarrow_\delta (\Pi x : A'. a')$$

$$\Gamma \vdash (\lambda x : A. a) \rightarrow_\delta (\lambda x : A'. a)$$

$$\Gamma \vdash A \rightarrow_\delta A'$$

$$\Gamma \vdash (ab) \rightarrow_\delta (ab')$$

$$\Gamma \vdash (\Pi x : A. a) \rightarrow_\delta (\Pi x : A'. a)$$

When $\Gamma$ is the empty context, it is written $a \rightarrow_\delta a'$ instead of $\Gamma \vdash a \rightarrow_\delta a'$.

The relation $\rightarrow_\beta$ between terms in $T$ will be extended to terms in $T_0$.

Definition A.3.
Let $\Gamma \in C_0$, $b, B \in T_0$. The DPTS determined by the specification $S = (S, A, R)$ is denoted as $\lambda S_\delta = \lambda (S, A, R)_\delta$ and defined by the notion of type derivation $\Gamma \vdash b : B$ given by the following axioms and rules:

$$(\text{axiom}) \quad \varepsilon \vdash c : s \quad \text{for } c : s \in A$$
\[
\begin{align*}
\text{(start)} & \quad \Gamma \vdash A : s \\
& \quad \Gamma, x : A \vdash x : A
\end{align*}
\]
where \( x \) is \( \Gamma \)-fresh

\[
\begin{align*}
\text{(weakening)} & \quad \Gamma \vdash b : B \quad \Gamma \vdash A : s \\
& \quad \Gamma, x : A \vdash b : B
\end{align*}
\]
where \( x \) is \( \Gamma \)-fresh

\[
\begin{align*}
\text{(formation)} & \quad \Gamma \vdash A : s_1 \quad \Gamma, x : A \vdash B : s_3 \\
& \quad \Gamma \vdash (\Pi x : A. B) : s_3
\end{align*}
\]
for \((s_1, s_2, s_3) \in R\)

\[
\begin{align*}
\text{(abstraction)} & \quad \Gamma, x : A \vdash b : B \quad \Gamma \vdash (\Pi x : A. B) : s \\
& \quad \Gamma \vdash (\lambda x : A. b) : (\Pi x : A. B)
\end{align*}
\]

\[
\begin{align*}
\text{(application)} & \quad \Gamma \vdash b : (\Pi x : A. B) \quad \Gamma \vdash a : A \\
& \quad \Gamma \vdash (b \ a) : B[x := a]
\end{align*}
\]

\[
\begin{align*}
\text{(conversion)} & \quad \Gamma \vdash b : B \quad \Gamma \vdash B' : s \quad B =_\beta B' \\
& \quad \Gamma \vdash b : B'
\end{align*}
\]

\[
\begin{align*}
\text{(\(\delta\) - start)} & \quad \Gamma \vdash a : A \\
& \quad \Gamma, z = a : A \vdash x : A
\end{align*}
\]
where \( x \) is \( \Gamma \)-fresh

\[
\begin{align*}
\text{(\(\delta\) - weakening)} & \quad \Gamma \vdash b : B \quad \Gamma \vdash a : A \\
& \quad \Gamma, z = a : A \vdash b : B
\end{align*}
\]
where \( x \) is \( \Gamma \)-fresh

\[
\begin{align*}
\text{(\(\delta\) - formation)} & \quad \Gamma, z = a : A \vdash B : s \\
& \quad \Gamma \vdash z = a : A \text{ in } B : s
\end{align*}
\]

\[
\begin{align*}
\text{(\(\delta\) - introduction)} & \quad \Gamma, z = a : A \vdash b : B \quad \Gamma \vdash x = a : A \text{ in } B : s \\
& \quad \Gamma \vdash x = a : A \text{ in } b : x = a : A \text{ in } B
\end{align*}
\]

\[
\begin{align*}
\text{(\(\delta\) - conversion)} & \quad \Gamma \vdash b : B \quad \Gamma \vdash B' : s \quad \Gamma \vdash B =_\beta B' \\
& \quad \Gamma \vdash b : B'
\end{align*}
\]

where \( s \) ranges over sorts, i.e. \( s \in S \).
Appendix B

Generalization for Abstract Reduction Systems

The notions of Abstract Reduction Systems (ARS) and sub-ARS are presented as in [Klo90].

Definition B.1.
An abstract reduction system (ARS) is a structure $\mathcal{A} = \langle A, (\rightarrow_\alpha)_{\alpha \in I} \rangle$ such that

- $A$ is a set
- For all $\alpha \in I$, $\rightarrow_\alpha \subseteq A \times A$. These binary relations are called reduction or rewrite relations.

Definition B.2. The relation $\rightarrow_\alpha$ is the transitive, reflexive closure of $\rightarrow_\alpha$.

Definition B.3. Let $\mathcal{A} = \langle A, \rightarrow_\alpha \rangle$ and $\mathcal{B} = \langle B, \rightarrow_\beta \rangle$ be two ARS’s. The concept "$\mathcal{A}$ is a sub-ARS of $\mathcal{B}$" is written as $\mathcal{A} \subset \mathcal{B}$ and defined as follows:

- $\mathcal{A} \subseteq \mathcal{B}$
- $\rightarrow_\alpha$ is the restriction of $\rightarrow_\beta$, i.e. $\forall a, a' \in A(a \rightarrow_\alpha a' \iff a \rightarrow_\beta a')$

$\mathcal{B}$ is also called an extension of $\mathcal{A}$.

Theorem B.4.
Let $\mathcal{A} = \langle A, \rightarrow_\alpha \rangle$ and $\mathcal{A}' = \langle A', \rightarrow_\beta \rangle$ be ARS’s be such that $\mathcal{A}'$ is an extension of $\mathcal{A}$. Let $\mathcal{T} = \langle A', \rightarrow_\tau \rangle$ be an ARS and suppose there is a mapping $\pi : A' \rightarrow A$ such that:

a) $\pi|A = id$

b) $\forall a \in A'(a \rightarrow_\beta \pi(a))$

c) If $a \rightarrow_\tau b$ then $\pi(a) = \pi(b)$ for all $a, b \in A'$.

d) If $a \rightarrow_\beta b$ then $\pi(a) \rightarrow_\beta \pi(b)$, for all $a, b \in A'$.

1. From b) and c) the following is deduced:
   $\beta$ is Church-Rosser

2. From b), c) and d) the following conclusion is deduced:
   If $\beta$ is Church-Rosser so is $\beta'b$. 

34
3. From a), b) and c) the following conclusion is deduced:

Suppose that $A$ is closed with respect to $\delta$ (i.e. $a \rightarrow_\delta b$ and $a \in A$ then $b \in A$) and that $\Delta = \{(d,d), d \rightarrow_\delta d\} = \emptyset$.

Then $\delta$ is weakly normalizing.

Proof:
1. The proof is shown in the following diagram.

\[
\begin{array}{c}
\text{a} \\
\delta \\
\text{a}_1 \quad \text{a}_2 \\
\delta \\
\pi(a_1) = \pi(a) = \pi(a_2)
\end{array}
\]

2. The proof is shown in the following diagram.

\[
\begin{array}{c}
\text{a} \\
\delta' \delta \\
\text{a}_1 \quad \text{a}_2 \\
\delta \\
\pi(a_1) = \pi(a) = \pi(a_2)
\end{array}
\]

3. Let be $a \in A'$. It follows from hypothesis b) that $a \rightarrow_\delta \pi(a)$. Suppose $\pi(a)$ is not in $\delta$-nf then $\pi(a) \rightarrow_\delta d$. Since $A$ is closed with respect to $\delta$, it follows that $d \in A$. It follows from hypothesis a) and c) that $\pi(a) = \pi(\pi(a)) = \pi(d) = d$. Then $d \rightarrow_\delta d$. Since $\Delta = \emptyset$, it follows that $\pi(a)$ is in $\delta$-nf.

Example B.5.
In the case of Chapter 4.

- $A_T \equiv \{a \in T, FV(a) \cap Def(\Gamma) = \emptyset\}, \neg \rho$
- $A_T' \equiv < T_\delta, \neg \rho >$
- $F_T \equiv < T_\delta, \neg \tau >$
- $\alpha \rightarrow_\delta b$ if and only if $\Gamma \vdash a \rightarrow_\delta b$
• \( \pi_r \equiv | - | r \)

Parts 1 and 2 of theorem B.4 are the generalization of the proofs in Section 4.2 for Church-Rosser for \( \rightarrow_\delta \) and for \( \rightarrow_\beta_\delta \). Part 3 of this theorem is somewhat different from the proof of weak normalization for \( \rightarrow_\delta \) presented in Section 4.3. In that section in order to prove that \( \pi(a) \) is in \( \delta \text{-nf} \), it is proved that all elements in \( \mathcal{A} \) are in \( \delta \text{-nf} \).

**Theorem B.6.** Let \( \mathcal{A} = \langle \mathcal{A}', \rightarrow_\beta \rangle, \mathcal{A}' = \langle \mathcal{A}', \rightarrow_\beta \rangle \) and \( \mathcal{F} = \langle \mathcal{A}', \rightarrow_\delta \rangle \) be ARS's. Suppose there is a mapping \( \pi : \mathcal{A}' \rightarrow \mathcal{A} \) such that:

a) If \( a \rightarrow_\delta b \) then \( \pi(a) \rightarrow_\beta \pi(b) \) for all \( a, b \in \mathcal{A}' \).

b) If \( a \rightarrow_\beta b \) then \( \pi(a) \rightarrow_\beta \pi(b) \) for all \( a, b \in \mathcal{A}' \).

If \( \delta \) and \( \beta \) are strongly normalizing. Then \( \beta' \delta \) is strongly normalizing.

**Proof:** Suppose towards a contradiction that \( \beta' \delta \) is not strongly normalizing, i.e. there is an infinite \( \beta' \delta \)-reduction sequence starting at \( a \in \mathcal{A}' \).

Observe that the number of \( \beta' \)-reductions in this sequence is infinite, i.e. \( \forall n \in \mathbb{N} \exists m > n \) : \( a_m \rightarrow_\delta a_{m+1} \). Otherwise it would follow that there is \( n_0 \in \mathbb{N} \) such that \( \forall m > n_0 \) \( a_m \rightarrow_\delta a_{m+1} \).

Hence the sequence \( a_{n+1} \rightarrow_\delta a_{n+2} \rightarrow \ldots \) would be infinite. As \( \delta \) is strongly normalizing, this can not happen. Hence the number of \( \beta' \)-reduction steps in the sequence \( a \rightarrow_\beta a_1 \rightarrow_\beta a_2 \ldots \) is infinite. Then this sequence is of the form

\[
a \rightarrow_\delta a_{n_1} \rightarrow_\beta a_{n_2} \rightarrow_\delta a_{n_3} \rightarrow_\beta a_{n_4} \rightarrow_\delta a_{n_5} \rightarrow_\beta a_{n_6} \rightarrow_\delta a_{n_7} \rightarrow \ldots
\]

By hypothesis a) and b) there is an infinite \( \beta \)-reduction sequence starting at \( \pi(a) \):

\[
\pi(a) \rightarrow_\delta \pi(a_{n_1}) \rightarrow_\beta \pi(a_{n_2}) \rightarrow_\delta \pi(a_{n_3}) \rightarrow_\beta \pi(a_{n_4}) \rightarrow_\delta \pi(a_{n_5}) \rightarrow_\beta \pi(a_{n_6}) \rightarrow_\delta \ldots
\]

which contradicts the assumption that \( \beta \) is strongly normalizing. \( \square \)

**Example B.7.**

In the case of Chapter 5, 

- \( \mathcal{A}_r \equiv \{ d \in T, \exists D (\{ \Gamma \} \vdash_D d : D) \} \) where \( S' \) is a completion of \( S \).
- \( \mathcal{A}_r^c = \{ d \in T, \exists D (\{ \Gamma \} \vdash_D d : D) \} \)
- \( \mathcal{A}_r^c = \langle \mathcal{A}_r, \rightarrow_\delta \rangle \)
- \( \mathcal{A}_r^c = \langle \mathcal{A}_r, \rightarrow_\delta \rangle \)
- \( \mathcal{F}_r \equiv \langle \mathcal{A}_r, \rightarrow_\delta \rangle \)
- \( a \rightarrow_\delta b \) if and only if \( \Gamma \vdash a \rightarrow_\delta b \)
- \( \pi_r \equiv \{ - \} r \)
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<table>
<thead>
<tr>
<th>Code</th>
<th>Author(s)</th>
<th>Title</th>
</tr>
</thead>
<tbody>
<tr>
<td>91/18</td>
<td>Rik van Geldrop</td>
<td>Transformational Query Solving, p. 35.</td>
</tr>
<tr>
<td>91/19</td>
<td>Erik Poll</td>
<td>Some categorical properties for a model for second order lambda calculus with subtyping, p. 21.</td>
</tr>
<tr>
<td>91/23</td>
<td>K.M. van Hee, L.J. Somers, M. Voorhoeve</td>
<td>Z and high level Petri nets, p. 16.</td>
</tr>
<tr>
<td>91/24</td>
<td>A.T.M. Aerts, D. de Reus</td>
<td>Formal semantics for BRM with examples, p. 25.</td>
</tr>
<tr>
<td>91/25</td>
<td>P. Zhou, J. Hooman, R. Kuiper</td>
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</tr>
<tr>
<td>91/27</td>
<td>F. de Boer, C. Palamidessi</td>
<td>Embedding as a tool for language comparison: On the CSP hierarchy, p. 17.</td>
</tr>
<tr>
<td>91/28</td>
<td>F. de Boer</td>
<td>A compositional proof system for dynamic process creation, p. 24.</td>
</tr>
<tr>
<td>91/30</td>
<td>J.C.M. Baeten, F.W. Vaandrager</td>
<td>An Algebra for Process Creation, p. 29.</td>
</tr>
<tr>
<td>91/31</td>
<td>H. ten Eikelder</td>
<td>Some algorithms to decide the equivalence of recursive types, p. 26.</td>
</tr>
<tr>
<td>91/33</td>
<td>W. v.d. Aalst</td>
<td>The modelling and analysis of queueing systems with QNM-ExSpect, p. 23.</td>
</tr>
<tr>
<td>91/34</td>
<td>J. Coenen</td>
<td>Specifying fault tolerant programs in deontic logic, p. 15.</td>
</tr>
</tbody>
</table>
92/01 J. Coenen
J. Zwiers
W.-P. de Roever

A note on compositional refinement, p. 27.

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A compositional semantics for fault tolerant real-time systems, p. 18.

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J.A. Bergstra

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F. Kamareddine

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92/20 F.Kamareddine

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92/21 F.Kamareddine

Non well-foundedness and type freeness can unify the interpretation of functional application, p. 16.
A useful lambda notation, p. 17.

Nominalization, Predication and Type Containment, p. 40.

Bottom-up Abstract Interpretation of Logic Programs, p. 33.

A Programming Logic for Fo, p. 15.

A modelling method using MOovie and SimCon/ExSpect, p. 15.

A taxonomy of keyword pattern matching algorithms, p. 50.

Deriving the Aho-Corasick algorithms: a case study into the synergy of programming methods, p. 36.

A continuous version of the Prisoner’s Dilemma, p. 17

Quicksort for linked lists, p. 8.

Deterministic and randomized local search, p. 78.

A congruence theorem for structured operational semantics with predicates, p. 18.

On the unavoidability of metastable behaviour, p. 29

Exercises in Multiprogramming, p. 97

A Formal Deterministic Scheduling Model for Hard Real-Time Executions in DEDOS, p. 32.


Systems Engineering: a Formal Approach Part II: Frameworks, p. 44.


<table>
<thead>
<tr>
<th>Ref.</th>
<th>Author(s)</th>
<th>Title and Page(s)</th>
</tr>
</thead>
<tbody>
<tr>
<td>92/22</td>
<td>R. Nederpelt, F. Kamareddine</td>
<td>A useful lambda notation, p. 17.</td>
</tr>
<tr>
<td>92/23</td>
<td>F. Kamareddine, E. Klein</td>
<td>Nominalization, Predication and Type Containment, p. 40.</td>
</tr>
<tr>
<td>92/24</td>
<td>M. Codish, D. Dams, Eyal Yardeni</td>
<td>Bottom-up Abstract Interpretation of Logic Programs, p. 33.</td>
</tr>
<tr>
<td>92/25</td>
<td>E. Poll</td>
<td>A Programming Logic for Fo, p. 15.</td>
</tr>
<tr>
<td>93/01</td>
<td>R. van Geldrop</td>
<td>Deriving the Aho-Corasick algorithms: a case study into the synergy of programming methods, p. 36.</td>
</tr>
<tr>
<td>93/02</td>
<td>T. Verhoeff</td>
<td>A continuous version of the Prisoner's Dilemma, p. 17</td>
</tr>
<tr>
<td>93/03</td>
<td>T. Verhoeff</td>
<td>Quicksort for linked lists, p. 8.</td>
</tr>
<tr>
<td>93/04</td>
<td>E. H. L. Aarts, J. H. M. Korst, P. J. Zwietering</td>
<td>Deterministic and randomized local search, p. 78.</td>
</tr>
<tr>
<td>93/05</td>
<td>J. C. M. Baeten, C. Verhoeff</td>
<td>A congruence theorem for structured operational semantics with predicates, p. 18.</td>
</tr>
<tr>
<td>93/06</td>
<td>J. P. Veltkamp</td>
<td>On the unavoidability of metastable behaviour, p. 29</td>
</tr>
<tr>
<td>93/07</td>
<td>P. D. Moerland</td>
<td>Exercises in Multiprogramming, p. 97</td>
</tr>
<tr>
<td>93/08</td>
<td>J. Verhoosel</td>
<td>A Formal Deterministic Scheduling Model for Hard Real-Time Executions in DEDOS, p. 32.</td>
</tr>
<tr>
<td>93/10</td>
<td>K. M. van Hee</td>
<td>Systems Engineering: a Formal Approach Part II: Frameworks, p. 44.</td>
</tr>
</tbody>
</table>

A Trace-Based Compositional Proof Theory for Fault Tolerant Distributed Systems, p. 27.

Hard Real-Time Reliable Multicast in the DEDOS system, p. 19.

A congruence theorem for structured operational semantics with predicates and negative premises, p. 22.

The Design of an Online Help Facility for ExSpect, p. 21.


A Typechecker for Bijective Pure Type Systems, p. 28.

Relational Algebra and Equational Proofs, p. 23.

Pure Type Systems with Definitions, p. 38.


Multi-dimensional Petri nets, p. 25.

Finding all minimal separators of a graph, p. 11.

A Semantics for a fine λ-calculus with de Bruijn indices, p. 49.

GOLD, a Graph Oriented Language for Databases, p. 42.

On Vertex Ranking for Permutation and Other Graphs, p. 11.


ILIAS, a sequential language for parallel matrix computations, p. 20.