2. Red. Systems & Term Rewriting

(a) Reductions in Agda.
(b) Reduction systems.
(c) Termination, confluence, normalisation.
(d) Term rewriting systems.

Notations in Agda

data N : Set where
  Z : N
  S : N → N

- data N : Set means that we have introduced a new set, which is given by the constructors which follow after the symbol where.
- What is in most programming languages called type is in Agda for historic reasons called “Set”.
- The above code introduces a new set, namely the set of natural numbers.
- It has two constructors: the constant Z, and S which takes as argument an n : N and returns an element of N.

N as a Reduction System

- So the elements of N are
  \[ Z \quad S \quad Z \quad S \quad (S \quad Z) \quad S \quad (S \quad (S \quad Z)) \quad \cdots \]
- We can now define + and * in N by induction over the definition of N.
- For those with mathematical problems: “Induction over the definition of N” means roughly case distinction on N in a terminating way.

(a) Reductions in Agda

- Functional programming is essentially based on term reduction:
- Assume we introduce the natural numbers as an algebraic data type built from 0 and S (this is actual Agda code):

  data N : Set where
  Z : N
  S : N → N

(reductionSystems1.agda)

- We write here Z instead of 0, since the symbol 0 will be reserved for the built-in integers.
- S n stands for n + 1.
Definition of +

\[ _+_: \mathbb{N} \to \mathbb{N} \to \mathbb{N} \]

\[ n + Z = n \]

\[ n + S m = S (n + m) \]

\[ _+_: \text{means that} \]

\[ \text{the first argument (denoted by the first \_)} \text{ of + is placed before +,} \]

\[ \text{the second argument (denoted by the second \_)} \text{ of + is placed after +,} \]

\[ \text{which means here that + is used infix:} \]

\[ \text{we write } s + t \text{ instead of } _+)_s t_. \]

Mixed Fixed Symbols

Agda allows arbitrary mixed fixed symbols:

\[ \text{For instance we can define } _\text{strange}_\text{symbol}_ \text{ as a symbol which is used as} \]

\[ n \text{ strange } m \text{ symbol } k \]

\[ \text{for} \]

\[ _\text{strange}_\text{symbol}_ n m k. \]

There are almost no restrictions.

If the parsing of an expression is ambiguous, a parse error is given – then one needs to resolve the ambiguity by using parentheses.

Definition of +

The definition of + above means that we have the following reductions:

\[ s + Z \rightarrow s \]

\[ s + S t \rightarrow S (s + t) \]

Note that \( S \) binds more than \( + \). So

\[ S r + s \text{ reads } (S r) + s. \]

\[ r + S s \text{ reads } r + (S s). \]

We have \( 2 + 2 \rightarrow 4: \)

\[ S (S Z) + S (S Z) \rightarrow S (S (S Z) + S Z) \]

\[ \rightarrow S (S (S (S Z) + Z)) \]

\[ \rightarrow S (S (S (S Z))) \]

Our first Agda Example

We are going to show how we can deal with this example in Agda.
Installation

A substantially improved version of Agda called Agda2 has been released recently.
In this module Agda2 will be used, which has a completely different syntax from Agda1.

Currently the installation requires some work.
Easy to compile versions (which exist for Agda1) are in preparation, but have not been released yet.
For Agda1 there exists a 1-click-Windows-installer.
We hope that this problem will soon be solved.

Anton Setzer has installed Agda2 under Linux.
He is working on getting it installed under Windows.
Agda can be installed under Macintosh as well.

Installation of Agda

Instructions on how to install Agda under Linux and hopefully as well under Windows will soon be created and can then soon be found under http://www.cs.swan.ac.uk/~csetzer/othersoftware/agda2/agda2installation.html

The installation will provide an Emacs mode for Agda files.
If a file with extension .agda is loaded into Emacs, then this mode is invoked.

The source code for the examples given in this lecture will be available from the course home page (the names of the files are added in the notes, e.g. in the form reductionSystems1.agda.

Agda in the Linux Lab

Agda will be installed in the Linux lab.
When this is ready, follow the item “Getting started with Agda” on the home page of this module.
Please check whether the installation works.

Help System of Emacs

Note that Emacs has an excellent and well-written help system.
Includes a hypertext version of most features of the Emacs, search facilities, descriptions of all variables.
The help system is activated using Control-h plus an additional key stroke.
Emacs notation: C-h.
A quick tutorial, which introduces the help system starts when typing, after Emacs has been started, C-h t
Working in Agda

Once, Agda is installed, the above can be defined as follows:

- One opens in Emacs a file with extension “.agda”, e.g. “reductionSystems1.agda”.
- Emacs will switch into Agda mode.
- Code written needs to be part of a “module” (we will not discuss details of the module system in this lecture course.)
- We will create a module by typing in:

\[
\text{module reductionSystems1 where}
\]

Blanks around “:”

- Please note that there needs to be a blank around all “:”.
- Z: without a blank in between is considered by Agda as an identifier Z:.
- :N without a blank in between is considered by Agda as an identifier :N.
- Only brackets “(”, “{”, “)” and blanks (and possibly some other symbols not discovered yet by A. Setzer) break identifiers.

Typing in Special Symbols

- The standard installation of Agda activates a special mode of Emacs which allows to type in special symbols.
- Special symbols are typed in by using command sequences inspired by \LaTeX{}.
- For instance
  - N is written by typing in \texttt{\Bbb{N}},
  - α by typing in \texttt{\alpha}.
- On the homepage for this course under “Other Course Material” a file leimListOfSymbols.tut (which was extracted from the source code of this mode) will be made available, which contains the key bindings for this mode.

Working in Agda

module reductionSystems1 where

Now we add the definition of N:

\[
\text{module reductionSystems1 where}
\]

\[
\text{data N : Set where}
\]

\[
\begin{align*}
\text{Z} & : \text{N} \\
\text{S} & : \text{N} \rightarrow \text{N}
\end{align*}
\]
Loading the Buffer

- Agda doesn’t realise any changes in the buffer, unless we load it.
- For this we can use the main menu, which one obtains by right-clicking on the word “Agda2” in the Panel.
- By choosing from the main menu “Load”, we load the buffer.
- An additional buffer called ∗All Goals∗ appears, which will be explained later.
- If there was an error loading the buffer, then instead of ∗All Goals∗ a buffer ∗Error∗ is displayed showing an error message.

Keyboard Short Cuts

- In the main menu, for each command a keyboard short cut is presented. It is advisable to learn the most frequently used ones.
- In order to type check the buffer, we can use the keyboard command C-c C-x C-l.

Goals

- When defining code, one can leave some code open for being filled in a later step.
- These holes are called goal, which stands for a term not yet defined.
- Syntax in Agda: {! !}, written in “green” in the Emacs mode.
- One can type in as well “?” for a goal, which will then be converted, when loading the buffer, into the symbol {! !}.
- So let’s type in the beginning of the definition of +:

\[
\_
+_\_ : \mathbb{N} \to \mathbb{N} \to \mathbb{N}
\]
\[
n + m = ?
\]
Goals

- Goals are numbered by the order in which they were created.
- Goals are displayed together with their type in separate buffer called “∗ Goals ∗”.
- This can be activated as well by using menu “Show Goals (C-c C-?)”.
- In Emacs mode, goals have a special status.
- When typing in text into a goal, the goal expands.

Using the goal-menu, we can find out what type is expected:

**Goal Type (C-c C-t).**
Agda shows: ?0 : N

There is as well a variant

**Goal Type (normalised).**
It evaluates the type of the goal using reduction rules (see later).

Using the main menu we can always show the types of all goals (Show Goals (C-c C-?)�).

We can as well jump to the next and previous goal using menu items Next Goal (C-c C-f) and Previous Goal (C-c C-p).

Case Distinction

\[- + - : \mathbb{N} \to \mathbb{N} \to \mathbb{N}\]

\[n + m = \{! !\}\]

In order to define \(n + m\), we have to make a case distinction on whether \(m = Z\) or \(m = S m'\).

This can be achieved by replacing the line \(n + m = \{! !\}\) by two lines for the two cases.

In order to achieve this we deactivate Agda by using main menu command **Deactivate Agda (C-c C-x C-d).**

Then we can replace the line \(n + m = \{! !\}\) by two lines as follows:

\[- + - : \mathbb{N} \to \mathbb{N} \to \mathbb{N}\]
\[n + Z = \{! !\}\]
\[n + S m = \{! !\}\]
Coverage Checker

- Agda has built in a **coverage checker**, which makes sure that if one makes a case distinction as above, then all cases are **covered**.
- If we omit one of the cases, e.g. the S-case, and load the buffer:

\[ _+_ : N \rightarrow N \rightarrow N \]
\[ n + Z = \{! !\} \]

then we get an error message (see next slide)

---

Give and Refine

\[ _+ _ : N \rightarrow N \rightarrow N \]
\[ n + Z = \{! !\} \]
\[ n + S m = \{! !\} \]

- We can solve the first goal by typing in the value \( n \).
- Then we can right-click on the goal and use from the goal-menu either **Give (C-c C-SPC)** or **Refine (C-d C-r)**.
  - Give works when one has an exact solution as in the situation above.
  - Refine works not always, but allows as well partial solutions, which need refinement. See the case \( S n \) on the next slide.

---

Error Message

```
/home/csetzer/lectures/07/
intertheo/agdalectureexamples/
reductionSystems1.agda:9,1-17 Incomplete pattern matching for _+_.
Missing cases:
_+_S_ when checking the definition of _+_.
```

- So in the definition of \( _+ _ \), the case where the first argument is arbitrary, and the second argument is of the form \( S \) applied to something is missing.

---

Refine

- We obtain in both cases

\[ _+ _ : N \rightarrow N \rightarrow N \]
\[ n + Z = n \]
\[ n + S m = \{! !\} \]

- We can use the refine mechanism in case of \( n + S m \) as follows:
  - We know that the solution will be of the form \( S \{! !\} \).
  - We can now type into the goal \( S \) and then use the command refine.
  - Agda knows that if we apply \( S \) to one argument (which is a natural number), then we get something which solves the goal.
Refine

\[+_+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N} \]
\[n + Z = n\]
\[n + S m = \{! !\}\]

- If we type into the goal \( S \) and refine it, we obtain

\[+_+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N} \]
\[n + Z = n\]
\[n + S m = S \{! !\}\]

Refine

Termination Checker

\[+_+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N} \]
\[n + Z = n\]
\[n + S m = S \{! !\} + \{! !\}\]

- The \(+\) and the defining symbol \(_+ _+\) are now marked in red.
- This is because of the termination checker.
- Agda disallows non-terminating programs, like

\[f : \mathbb{N} \to \mathbb{N} \]
\[f n = \]
Termination Checker

In the code

\[
\begin{align*}
\text{\_ + \_} &: \mathbb{N} \to \mathbb{N} \to \mathbb{N} \\
n + Z &= n \\
n + S \, m &= S \, (\{! \} \oplus \{! \!\!}
\end{align*}
\]

we don’t know yet whether this will pass the termination checker when the goals are solved or not.

If we solve it by

\[
n + S \, m = S \, (n \oplus S \, m)
\]

then we obtain a non-terminating program.

---

Termination Checker

If we solve

\[
\begin{align*}
\text{\_ + \_} &: \mathbb{N} \to \mathbb{N} \to \mathbb{N} \\
n + Z &= n \\
n + S \, m &= S \, (\{! \} \oplus \{! \!\!}
\end{align*}
\]

in the correct way

\[
n + S \, m = S \, (n + m)
\]

it will pass the termination checker.

---

Termination Checker

Note that only a warning (in the form of the symbols coloured) is issued, but no error is issued.

This is since this warning only indicates that there is potential problem.

It might be solved once all goals are solved.

There are as well limitations to any termination checker:

It is in principal not possible to write a termination checker which accepts all terminating Agda programs.

So a program could be terminating and therefore okay, but still not pass the termination checker.

This will be discussed later.

---

Termination Checker

In order to make sure that there are no termination check problems left, one can use from a shell the command

```
agda file
```
e.g.

```
agda ~csetzer/r/reductionSystems1.agda
```

This command will check the file and report type errors, problems of the coverage checker and problems of the termination checker.

**Code submitted as coursework should be checked this way**, in order to guarantee that there are no hidden errors.
Finishing Definition of +

\[+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N}\]
\[n + Z = n\]
\[n + S \, m = S \, (n + m)\]

We can solve now the two goals by using \(n\) and \(m\) and
goal-menu refine or give and obtain

\[+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N}\]
\[n + Z = n\]
\[n + S \, m = S \, (n + m)\]

Indentation Sensitivity

Agda is indentation sensitive.

So often instead of having parentheses “\(\{\text{Code}\}\)”, as
in other languages, all lines belonging to \(\langle\text{Code}\rangle\) have to
be intended more then the surrounding code, and
usually in the same way.

Therefore top level definitions have to start in column 1.
Otherwise they are considered as being an extension of
a previous definition.

All code belonging to such a definition in later columns
has to be intended at least once.

Example: The following causes an error:

\[
\text{data } N : \text{Set where}
\]
\[Z : N\]
\[S : N \to N\]

Agda assumes that \(S\) is not a constructor of \(N\), but a
function, which is not defined yet.

Example: The following causes an error:

\[
\text{data } N : \text{Set where}
\]
\[Z : N\]
\[S : N \to N\]

\[+ : \mathbb{N} \to \mathbb{N} \to \mathbb{N}\]
\[n + Z = n\]
\[n + S \, m = S \, (n + m)\]

Now \(+\) is considered as a constructor of \(N\), and the
equation \(n + Z = n\) doesn't make sense for a
constructor and causes a parse error.
Definition of Multiplication

Definition of \( \_ \ast \_ \):

\[
\_ \ast \_ : \mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N}
\]

\[
n \ast Z = Z
\]

\[
n \ast S \, m = (n \ast m) + n
\]

This means that we have the following reductions:

\[
s \ast Z \rightarrow Z,
\]

\[
s \ast S \, t \rightarrow s \ast t + s.
\]

Binding of Symbols

We can add the following two lines:

\[
infixl 60 \_ + \_
\]

\[
infixl 80 \_ \ast \_
\]

This means that \( \_ + \_ \) and \( \_ \ast \_ \) are infix left-associative:

\( n + m + k \) is interpreted as \((n + m) + k\), similarly for \( \ast \).

If we had used \( infixr \) instead, we obtain

\( n + m + k \) is interpreted as \( n + (m + k) \), similarly for \( \ast \).

If we use \( infix \), then \( n + m + k \) is considered as ambiguous and causes a parse error.

Complete Definition of \( + \) and \( \ast \)

We obtain as definition of \( + \) and \( \ast \):

\[
infixl 60 \_ + \_
\]

\[
infixl 80 \_ \ast \_
\]

Furthermore that \( \_ \ast \_ \) has a higher number than \( \_ + \_ \) means that \( \ast \) binds more than \( + \):

\( n + m \ast k \) is interpreted as \( n + (m \ast k) \).

Without stating the statements above \( n + m \ast k \) is considered as ambiguous and causes a parse error.
Testing the above in Agda

- In order to test the above we can make use of the main-menu (or goal-menu)
  Compute normal form (C-c C-n).
- It will ask in the mini-buffer for an expression.
- If we type in
  \[ S (S Z) + S (S Z) \]
  (for \( 2 + 2 \)), Agda shows in another buffer the result
  \[ S (S (S (S Z))) \]
  i.e. \( 4 \).

Using the Builtin Natural Numbers

- Then we can define for instance
  \[
  a : \mathbb{N} \\
  a = 5
  \]
- and if we compute the normal form of \( 7 + 9 \) we obtain \( 16 \).

Using the Builtin Natural Numbers

- We can use as well the builtin natural numbers:
  If we add the following code
  \[
  \{- \# \text{ BUILTIN NATURAL } \mathbb{N} \#-\} \\
  \{- \# \text{ BUILTIN ZERO } Z \#-\} \\
  \{- \# \text{ BUILTIN SUC } S \#-\} \\
  \{- \# \text{ BUILTIN NATPLUS } _+_- \#-\} \\
  \{- \# \text{ BUILTIN NATTIMES } _*__- \#-\}
  \]
  then \( \mathbb{N} \) is identified with the builtin type of natural numbers, and \( Z, S, _+_, _*_ \) with the corresponding builtin operations.

Agda and Non-Termination

- If one makes a mistake and defines \( _+_ \) so that it doesn’t terminate (e.g. defining in case of \( S m \)
  \( S (n + S m) \) instead of \( S (n + m) \)), then Agda will crash, and not display anything.
  This can be observed by checking the buffer \(*ghci*\).
  All Emacs activities will result in Haskell commands been issued to this buffer, and the result is then used in order to modify the emacs buffer.
  Since the interactive Glasgow Haskell Compiler ghci is used, the buffer for communicating is called \(*ghci*\).
  If Agda crashes, one sees that a command was issued there, but no response was returned yet.
Agda and Non-Termination

- If one has non-terminating recursion, Agda might crash during type checking and at other places as well.
- Can be observed as well by switching to buffer `*ghci*`.
- The ghci buffer is a buffer to Haskell with Agda loaded.
- It can be used for carrying out haskell computations, e.g. for computing $3 + 3$ in Haskell.

Towards General Reduction Systems

- Since we have dependent types, equality plays a rôle in type checking
  - If $A r$ is a type depending on $r$ and $a : A r$, then $a : A s$ provided that $r$ and $s$ are equal.
- In order to understand Agda better, we will study in the following general reduction relations.
- Given by a set of Terms $T$ and a reduction relation $s \rightarrow t$ between terms $s$ and $t$.

Reduction Systems and Agda

- We want $2 + 2$ and $4$ to be the same.
- In Agda, referring to our self-defined natural numbers, this means that $S (S Z) + S (S Z)$ and $S (S (S (S Z)))$ should be the same.
- We have just seen that $S (S Z) + S (S Z)$ reduces to $S (S (S (S Z)))$.
- The underlying principle behind this is:
  
  If a term reduces to another term, then these two terms are the same.

(b) Reduction Systems

- A reduction system is a pair $(T, \rightarrow)$ consisting of a set $T$ (of terms) and a binary relation $\rightarrow$ on $T$.
- We write $s \rightarrow t$ for “$s, t$ are in relation $\rightarrow$” and say usually “$s$ reduces to $t$”.
- Example 1:
  - Let $T$ be the set of terms formed from $0$, $S$, $+$ and $\times$ in the usual way.
  - So for instance $0$, $S 0$ and $S 0 + 0$ are elements of $T$.
  - Let $\rightarrow$ be the reduction relation defined as before.
  - So we have for instance
    
    $S(S 0) + 0 \rightarrow S(S 0)$
    $S(S 0) + S 0 \rightarrow S(S(S 0))$
    
    Then $(T, \rightarrow)$ forms a reduction system.
Example 2 (Reduction System)

A simple reduction system is $T = \mathbb{N}$ with reductions $n + 1 \rightarrow n$ for $n \in \mathbb{N}$:

$0 \leftarrow 1 \leftarrow 2 \leftarrow 3 \leftarrow 4 \leftarrow \cdots$

So we have reductions of the form:

$5 \rightarrow 4 \rightarrow 3 \rightarrow 2 \rightarrow 1 \rightarrow 0$.

Example 3 (Reduction System)

Another simple example is $T = \mathbb{N}$ with reductions $n \rightarrow m$ for $n, m \in \mathbb{N}$ s.t. $n > m$:

$0 \leftarrow 1 \leftarrow 2 \leftarrow 3 \leftarrow 4$

So we have reductions of the form:

$23 \rightarrow 11 \rightarrow 3 \rightarrow 1 \rightarrow 0$.

Example 4 (Reduction System)

A further simple example is $T = \mathbb{N} \cup \{\bullet\}$, with reductions $n + 1 \rightarrow n$ for $n \in \mathbb{N}$, and $\bullet \rightarrow n$ for $n \in \mathbb{N}$:

So we have reductions of the form:

$\bullet \rightarrow 5 \rightarrow 4 \rightarrow 3 \rightarrow 2 \rightarrow 1 \rightarrow 0$.

If $(T, \rightarrow)$ is a reduction system, we define for $s, t \in T$

$s \leftarrow t \iff t \rightarrow s$

$s \leftarrow t \iff s \rightarrow t \lor s \leftarrow t$

Note that we are using “$\lor$”, not “$\land$”.
In Agda we said we identify two terms which reduce to each other in possible multiple steps.

Therefore we study two concepts:
- One is $s \rightarrow^{\ast} t$, which means that $s$ reduces to $t$ in possibly multiple steps.
- When Agda reduces a term $s$, it returns a term $t$ s.t. $s \rightarrow^{\ast} t$, and $t$ cannot reduce any further.
- One is $s \leftrightarrow^{\ast} t$, which is the equality induced by $\rightarrow$.
- So Agda identifies terms $s$ and $t$ s.t. $s \leftrightarrow^{\ast} t$.

---

If $(T, \rightarrow)$ is a reduction system, we define
- $s \equiv^{\ast} t$ iff there exists a (possibly empty) sequence
  
  \[ s \equiv s_0 \rightarrow s_1 \rightarrow s_2 \rightarrow \cdots \rightarrow s_n \equiv t \]

  By empty sequence we mean: $n = 0$ is allowed, in which case we have $s \equiv t$.

  (We write $\equiv$ for syntactic equality between terms in order to avoid confusion with equality modulo reductions introduced later and denoted by $\equiv$).

---

In order to express the above shorter, one says that $\rightarrow^{\ast}$ is the reflexive and transitive closure of $\rightarrow$, i.e. the least reflexive and transitive relation containing $\rightarrow$.

This means the following:
- $r \rightarrow s$ implies $r \rightarrow^{\ast} s$.
- $\rightarrow^{\ast}$ is reflexive, i.e. for all $r \in T$ we have $r \rightarrow^{\ast} r$.
- $\rightarrow^{\ast}$ is transitive, i.e. $r \rightarrow^{\ast} s \rightarrow^{\ast} t$ implies $r \rightarrow^{\ast} t$.
- If there is any other relation $\rightarrow'$, which is reflexive, transitive and contains $\rightarrow$, then $r \rightarrow^{\ast} s$ implies $r \rightarrow' s$.

The next slides contain a proof that $\rightarrow^{\ast}$ is indeed the reflexive transitive closure of $\rightarrow$. We jump over it. Jump over proof.

---

Example

If we take $\mathbb{N}$ with reductions $n + 1 \rightarrow n$ for $n \in \mathbb{N}$:

\[ 0 \leftrightarrow 1 \leftrightarrow 2 \leftrightarrow 3 \leftrightarrow 4 \leftrightarrow \cdots \]

Then $5 \rightarrow 4 \rightarrow 3 \rightarrow 2$, therefore $5 \rightarrow^{\ast} 2$.

In general $n \rightarrow^{\ast} m \leftrightarrow n \geq m$.
**Proof**

We show that \( \rightarrow^* \), as defined originally, is in fact the reflexive and transitive closure of \( \rightarrow \):

- If \( r \rightarrow s \), then clearly \( r \rightarrow^* s \) (take \( n = 1 \)).
- \( \rightarrow^* \) is reflexive: We have \( r \rightarrow r \) for \( r \in T \), by having \( n = 0 \) in the definition of \( \rightarrow^* \).
- \( \rightarrow^* \) is transitive: Assume \( r \rightarrow^* s \rightarrow^* t \). Then there exist \( n, m, r_i, s_i \) s.t.
  \[
  r \equiv r_0 \rightarrow r_1 \rightarrow \cdots \rightarrow r_n \equiv s \\
  s \equiv s_0 \rightarrow s_1 \rightarrow \cdots \rightarrow s_m \equiv t 
  \]

But this implies \( r \rightarrow^* t \).

**Proof (Cont.)**

Furthermore, if \( \rightarrow' \) is another relation, which contains \( \rightarrow \), and which is reflexive and transitive, then it contains \( \rightarrow^* \):

- Assume \( \rightarrow' \) having these properties.
- Assume \( r \rightarrow^* s \).
- Then there exist \( n \in \mathbb{N}, r_i \) s.t.
  \[
  r \equiv r_0 \rightarrow r_1 \rightarrow \cdots \rightarrow r_n \equiv s 
  \]

In case \( n = 0 \) we have \( r \equiv s \), therefore \( r \rightarrow' s \).

**Proof (Cont.)**

If \( (T, \rightarrow) \) is a reduction system, we define \( \leftarrow^* \):

- \( s \leftarrow^* t \) iff there exists a (possibly empty) sequence
  \[
  s \equiv s_0 \leftarrow s_1 \leftarrow s_2 \leftarrow \cdots \leftarrow s_n \equiv t 
  \]

In case \( n > 0 \) we have by the fact that \( t \rightarrow t' \) implies \( t \rightarrow' t' \) that

\[
  r \equiv r_0 \rightarrow r_1 \rightarrow \cdots \rightarrow r_n \equiv s 
  \]

But since \( \rightarrow' \) is transitive, it follows

\[
  r \rightarrow' s 
  \]

and we are done.

Note how easy it is to overlook that the case \( n = 0 \) has to be treated separately.

If one does this more formally, or using an interactive theorem proving, one would notice this missing case.
\[ r \longrightarrow^* t \iff s \equiv s_0 \longleftrightarrow s_1 \longleftrightarrow s_2 \longleftrightarrow \cdots \longleftrightarrow s_n \equiv t \]

- Note that if we want to identify two elements \( r, s \), if \( r \longrightarrow s \), we have to identify \( r, s \) if \( r \longrightarrow^* s \).
- If we want to identify elements \( r, s \) s.t. \( r \longrightarrow s \), we have to identify as well elements \( r, s \) s.t. \( r \longleftrightarrow s \).
- Then we have to identify elements \( r, s \) s.t. \( r \longrightarrow s \).
- Therefore we have, if \( n, s_i \) are as in the definition of \( s \longrightarrow^* t \), to identify \( s_0 \) and \( s_1; s_1 \) and \( s_2; \) etc.; \( s_{n-1} \) and \( s_n \).
- Therefore we have to identify \( s_0 \) and \( s_n \).

### Example

Assume the following reduction system:

\[ \begin{array}{ccc}
0 & \rightarrow & 3 \\
4 & \rightarrow & 1 \\
& \rightarrow & 2
\end{array} \]

Then \( 0 \leftarrow 3 \leftarrow 1 \leftarrow 4 \leftarrow 2 \), therefore \( 0 \leftarrow^* 2 \).

Jump over next slide.

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In order to express the above shorter, one says that \( \longleftrightarrow^* \) is the reflexive, symmetric and transitive closure of \( \longrightarrow \), i.e. the least reflexive, symmetric and transitive relation containing \( \longrightarrow \).

- This means the following:
  - \( s \longrightarrow t \) implies \( s \longleftrightarrow^* t \).
  - \( \longleftrightarrow^* \) is reflexive, and transitive.
  - \( \longleftrightarrow^* \) is symmetric, i.e. \( s \longleftrightarrow^* t \) implies \( t \longleftrightarrow^* s \).
  - If there is any relation \( \longleftrightarrow' \) with the above properties, then \( s \longleftrightarrow^* t \) implies \( s \longleftrightarrow' t \).
- The next slides contain a proof that \( \longleftrightarrow^* \) is indeed the reflexive, symmetric and transitive closure of \( \longleftrightarrow \). We jump over it. Jump over proof.
Proof

We show that $\rightsquigarrow^*$, as defined originally, is in fact the reflexive, symmetric, and transitive closure of $\rightsquigarrow$:

- That it contains $\rightsquigarrow$ follows since $r \rightsquigarrow s$ implies $r \rightsquigarrow s$ by an argument similar to that for $\rightsquigarrow^*$.
- That it is reflexive and transitive follows as for $\rightsquigarrow^*$.
- $\rightsquigarrow^*$ is symmetric: Assume $r \rightsquigarrow^* s$. Then there exist $n, r_i$, s.t.
  \[ r \equiv r_0 \rightsquigarrow r_1 \rightsquigarrow \cdots \rightsquigarrow r_n \equiv s \]

Now $r_i \rightsquigarrow r_{i+1}$ implies $r_{i+1} \rightsquigarrow r_i$, therefore

\[ s \equiv r_n \rightsquigarrow r_{n-1} \rightsquigarrow \cdots \rightsquigarrow r_0 \equiv r \]

which implies $s \rightsquigarrow^* r$.

Proof (Cont.)

Furthermore, if $\rightsquigarrow'$ is another relation which contains $\rightsquigarrow$ and which is reflexive, symmetric and transitive then it contains $\rightsquigarrow^*$:

- Assume $\rightsquigarrow'$ having these properties.
- Assume $r \rightsquigarrow^* s$.
- Then there exist $n \in \mathbb{N}, r_i$ s.t.
  \[ r \equiv r_0 \rightsquigarrow r_1 \rightsquigarrow \cdots \rightsquigarrow r_n \equiv s \]

In case $n = 0$ we have $r \equiv s$, therefore $r \rightsquigarrow' s$.

In case $n > 0$ we first note that, since $\rightsquigarrow'$ contains $\rightsquigarrow$ and is symmetric, we have that $t \rightsquigarrow t'$ implies $t \rightsquigarrow' t'$ and $t' \rightsquigarrow t$ implies $t \rightsquigarrow' t'$.

Therefore $t \rightsquigarrow t'$ implies $t \rightsquigarrow' t'$, and we obtain by the above

\[ r \equiv r_0 \rightsquigarrow' r_1 \rightsquigarrow' \cdots \rightsquigarrow' r_n \equiv s \]

But since $\rightsquigarrow'$ is transitive, it follows

\[ r \rightsquigarrow' s \]

and we are done.

Identification of Elements

- If we have a reduction system $(T, \rightarrow)$, one writes
  \[ s \rightarrow t \]

  or sometimes
  \[ s = t \]

  for $s \rightsquigarrow^* t$.

- In order to avoid confusion, we write
  \[ s \equiv t \]

  for $s$ and $t$ are the same element of $T$ without using any reductions.
Determination of $\leftrightarrow^*$

In general it is infeasible to determine whether $s \leftrightarrow^* t$ holds.

One has to check all possible ways of getting from $s$ to $t$, by both using $\rightarrow$ and $\leftarrow$.

In many cases this can be determined by:

- Reducing $s$ to some term $s'$ s.t. $s \rightarrow^* s'$ and $s'$ has no further reductions,
  i.e. by “evaluating $s$”.
- Doing the same with $t$ to some term $t'$.
- Checking whether $s'$ is identical to $t'$.

This way of determining, whether $s \rightarrow^* t$ holds, is correct, if $\rightarrow$ is confluent and strongly normalising (see next subsection).

Strong Normalisation

A reduction system $(T, \rightarrow)$ is **terminating** or **strongly normalising**, iff there is no infinite sequence $s_0 \rightarrow s_1 \rightarrow s_2 \rightarrow s_3 \rightarrow \cdots$ of elements in $T$.

Examples

- The following reduction system is terminating:

  $\begin{array}{cccccc}
  0 & \leftarrow & 1 & \leftarrow & 2 & \leftarrow & 3 & \leftarrow & 4 & \leftarrow & \cdots
  \end{array}$

  Any reduction sequence will end in 0 and terminate.

- The following reduction system is non terminating:

  $\begin{array}{cccccc}
  0 & \rightarrow & 1 & \rightarrow & 2 & \rightarrow & 3 & \rightarrow & 4 & \rightarrow & \cdots
  \end{array}$

  (Take $0 \rightarrow 1 \rightarrow 2 \rightarrow 3 \rightarrow \cdots$).
Examples

The following reduction system is terminating, but there are arbitrarily long reduction sequences starting with •:

\[ 0 \leftarrow 1 \leftarrow 2 \leftarrow 3 \leftarrow 4 \leftarrow \ldots \]

We have \( \bullet \rightarrow n \rightarrow (n-1) \rightarrow (n-2) \rightarrow \cdots \rightarrow 0. \)

Examples

The untyped \( \lambda \)-calculus (See next Sect.) is non terminating, since we have for \( \Omega := (\lambda x.x \ x) \ (\lambda x.x \ x) \):

\( \Omega \rightarrow \Omega \rightarrow \Omega \rightarrow \cdots \)

The typed \( \lambda \)-calculus (see later Subsection) is terminating.

In fact, the typed \( \lambda \)-calculus was introduced in order to obtain a terminate subtheory of the \( \lambda \)-calculus.

Normal Form and Irreducibility

Let \((T, \rightarrow)\) be a reduction system.

- \( s \in T \) is **irreducible**, if there exists no \( t \in T \) s.t. \( s \rightarrow t. \)
- \( t \) is a **normal form of** \( s \) iff \( s \rightarrow^* t \) and \( t \) is irreducible.
- \((T, \rightarrow)\) is **weakly normalising** or **normalising**, if every \( s \in T \) has a normal form.

Example

The following system is weakly normalising, but not strongly normalising:

\[ 0 \rightarrow 1 \rightarrow 2 \rightarrow 3 \rightarrow 4 \rightarrow \cdots \]

Every \( r \) has a normal form, namely •.

But there exists an infinite sequence

\[ 0 \rightarrow 1 \rightarrow 2 \rightarrow 3 \rightarrow \cdots \]
Example 2

The following system is both weakly normalising and strongly normalising:

If one reduces any element as long as possible, one finally ends up with $\bullet$, which doesn’t reduce any further.

So every element has the same normal form namely $\bullet$.

Example 3

In the following system 0 has two normal forms, namely 1 and 2:

This system is both strongly and weakly normalising, but is not confluent. (“Confluent” will be defined later).

Lemma

Let $(T, \rightarrow)$ be a strongly normalising reduction system. Then $(T, \rightarrow)$ is weakly normalising.

Proof:
A normal form of $s \in T$ can be obtained by simply reducing $s$ as long as possible:
Since $(T, \rightarrow)$ is strongly normalising, the reduction sequence terminates in some $t \in T$.
$t$ is a normal form of $s$.

Church-Rosser

We say a reduction system $(T, \rightarrow)$ is confluent or has the Church-Rosser property iff for all $r, s, t \in T$ we have

- if $r \rightarrow^* s$ and $r \rightarrow^* t$,
- then there exists an $t'$ s.t. $s \rightarrow^* t'$ and $t \rightarrow^* t'$. 
**Diamond Property**

Because of the shape of the picture on the previous slide, the Church-Rosser property is sometimes called as well the

- **Diamond property** or
- **Triangle property**.

So Church-Rosser means:

*Every triangle (or better fork) can be closed to a diamond.*

---

**Weak Church-Rosser**

But that condition is weaker than full Church-Rosser.

The following term rewriting system is weakly Church-Rosser, but doesn’t fulfil the full Church-Rosser Property:

\[ r \rightarrow^* r', r \rightarrow^* s' \text{ but there is no } t \text{ s.t. } r' \rightarrow^* t \text{ and } s' \rightarrow^* t. \]

---

**Theorem**

If \((T, \rightarrow)\) is confluent, then we have for \(r, s \in T\):

- \(r \leftarrow \rightarrow^* s\)
- iff there exists a \(t \in T\) s.t.
  - \(r \rightarrow^* t \land s \rightarrow^* t\)

---

Before the slide change, Church-Rosser was described a bit differently. Now we see an example of a term rewriting system that is weakly Church-Rosser but doesn’t satisfy Church-Rosser. The theorem states that if a term rewriting system is confluent, then there exists a relationship between terms that satisfies a certain condition. This relationship is crucial for understanding the behavior of the system and its properties.
Idea of Proof

Proof of the Theorem

We define for \( r, s \in T \)

\[
\downarrow s \iff \exists t \in T. (r \rightarrow^* t \land s \rightarrow^* t)
\]

So \( \downarrow s \) means that \( r \) and \( s \) have a common reduct:

Proof of the Theorem

So we have to show

\[
r \leftrightarrow^* s \iff r \downarrow s
\]

“\( \iff \)” is easy. If \( r \rightarrow^* t, s \rightarrow^* t \), then we get

\[
r \rightarrow^* t \leftrightarrow^* s
\]

(where \( t \leftrightarrow^* s \iff s \rightarrow^* t \)) and therefore \( r \leftrightarrow^* s \)

Proof of the Theorem

For the more difficult direction “\( \Rightarrow \)” we give two proofs:

- One more concrete and intuitive one.
  - Will be presented during the lecture.
- One more abstract one.
  - Will not be presented during the lecture.
First Proof of "⇒"

Assume $r \leftrightarrow^* s$.

This means that we have a chain

$$r \equiv r_0 \leftrightarrow r_1 \leftrightarrow r_2 \leftrightarrow \cdots \leftrightarrow r_n \equiv s$$

We are going to show successively:

- $r_0 \downarrow r_0$,
- $r_0 \downarrow r_1$,
- $r_0 \downarrow r_2$,
- $\cdots$
- $r_0 \downarrow r_n \equiv s$ (the assertion).

Diagram

That $r \downarrow s \leftrightarrow t$ implies $r \downarrow t$ can be visualised as follows:

```
  r
  |  *
  v  *
  s  t
```

First Proof of "⇒"

$r \downarrow s \iff \exists t. (r \rightarrow^* t \land s \rightarrow^* t)$.

In order to show this we have to show the following:

- For the first step, we need to show $r_0 \downarrow r_0$, i.e. in general we need to show (1) $r \downarrow r$.

For the step from $r_0 \downarrow r_i$ to $r_0 \downarrow r_{i+1}$ we need to show:

- If $r_0 \downarrow r_i$ and $r_i \leftrightarrow r_{i+1}$ then $r_0 \downarrow r_{i+1}$.
- In general we have to show:

  If $r \downarrow s$ and $s \leftrightarrow t$, then $r \downarrow t$,

  in short:

  $r \downarrow s \leftrightarrow t$ implies $r \downarrow t$.

First Proof of "⇒"

(1) $r \downarrow r$.

(2) $r \downarrow s \leftrightarrow t$ implies $r \downarrow t$.

Since $s \leftrightarrow t$ means $s \rightarrow t$ or $s \leftarrow t$, we need to show (2) $r \downarrow s \rightarrow t \Rightarrow r \downarrow t$,

(3) $r \downarrow s \leftarrow t \Rightarrow r \downarrow t$.

So in total we have to show (1), (2), (3) above.
(1) \( r \downarrow r \)

We show \( r \downarrow r \).

Formally: We have \( r \rightarrow^* r \) and \( r \rightarrow^* r \), therefore \( r \downarrow r \).

(2) \( r \downarrow s \rightarrow t \) implies \( r \downarrow t \)

Assume \( r \downarrow s \), \( s \rightarrow t \). Show \( r \downarrow t \).

\( r \rightarrow^* r' \), \( s \rightarrow^* r' \) for some \( r' \).

By Church-Rosser, \( s \rightarrow^* r' \) and \( s \rightarrow^* t \) implies that there exists a \( s' \) s.t. \( r' \rightarrow^* s', \ t \rightarrow^* s' \).

But then \( r \rightarrow^* r' \rightarrow^* s' \) therefore \( r \rightarrow^* s' \),

\( t \rightarrow^* s' \),

therefore \( r \downarrow t \).

(3) \( r \downarrow s \leftarrow t \) implies \( r \downarrow t \)

Assume \( r \downarrow s \), \( s \leftarrow t \). Show \( r \downarrow t \).

\( r \rightarrow^* r' \), \( s \rightarrow^* r' \) for some \( r' \).

But then

\( r \rightarrow^* r' \)

\( t \rightarrow^* s \rightarrow^* r' \),

therefore \( r \downarrow t \).
$r \downarrow s \leftarrow t$ implies $r \downarrow t$

This completes the first proof of the Theorem.

Second Proof of "⇒"

We show that $\downarrow$ contains $\rightarrow$ and is reflexive, symmetric and transitive.

By definition, $\leftrightarrow$ is the least reflexive, symmetric and transitive relation which contains $\rightarrow$.

Therefore $\leftrightarrow$ is contained in $\downarrow$ and we obtain

$r \leftrightarrow^* s \Rightarrow r \downarrow s$

Formally: If $r \rightarrow s$ then we have with $t := s$ that $r \rightarrow^* t$ and $s \rightarrow^* t$.

$\downarrow$ is reflexive. (As in the previous proof).

Formally: We have $r \rightarrow^* r$ and $r \rightarrow^* r$, therefore $r \downarrow r$. 
Second Proof of “⇒”

- ↓ is symmetric:

![Diagram showing symmetry]

Formally:
Assume \( r \downarrow s \).
Then \( r \longrightarrow^* t \) and \( s \longrightarrow^* t \) for some \( t \).
Therefore \( s \downarrow r \).

Second Proof of “⇒”

- ↓ is transitive:

![Diagram showing transitivity]

Formally:
Assume \( r \downarrow s \) and \( s \downarrow t \).
Then there exists \( r', s' \) s.t. \( r \longrightarrow^* r' \), \( s \longrightarrow^* s' \), \( t \longrightarrow^* t' \).
Then by confluence there exists an \( t' \) s.t. \( r' \longrightarrow^* t' \), \( s' \longrightarrow^* t' \).
Then \( r \longrightarrow^* t' \) and \( t \longrightarrow^* t' \).
Therefore \( r \downarrow t \).

Unique Normal Forms

**Lemma:**
Let \((T, \longrightarrow)\) be a confluent reduction system.
If \( r \in T \) has a normal form \( s \), then it is unique:
- If \( t \) is another normal form, then \( s \equiv t \).

**Proof:**
- We have \( r \longrightarrow^* s \) and \( r \longrightarrow^* t \).
- By confluence, there exists a \( r' \) s.t. \( s \longrightarrow^* r' \) and \( t \longrightarrow^* r' \).
- But since \( s \) and \( t \) are normal forms, it follows \( s \equiv r' \) and \( t \equiv r' \).
Lemma

Let \((T, \rightarrow)\) be a weakly normalising and confluent reduction system. Then

\[ r \leftrightarrow^* s \text{ iff the normal forms of } r \text{ and } s \text{ coincide.} \]

Proof:

\(\Rightarrow\): By Church Rosser \(r \leftrightarrow^* s\) implies the existence of a \(t\) s.t. \(r \rightarrow^* t\) and \(s \rightarrow^* t\).

Reduce \(t\) further to a normal form \(r'\).

Then \(r'\) is a normal form of both \(r\) and \(s\) as well.

Since by the above lemma, normal forms are unique, \(r'\) is the normal form of \(r\) and \(s\).

Lemma

\(\Leftarrow\): If the normal forms \(t\) coincide, then we have \(r \rightarrow^* t \leftarrow^* s\), therefore \(r \leftrightarrow^* s\).
Remark on Agda

- The underlying reduction system of Agda is strongly normalising and confluent, provided the code has been termination checked.
- The equality derived from this reduction system is used in order to typecheck terms.

Example of a Term Rewriting System

- Take $T = \text{set of arithmetic expressions formed from variables, } 0 \text{ by using the successor operation } S$ (where $S n \text{ stands } n + 1$), $+, \times$ and brackets.
- So the following are elements of $T$:
  - $x + S 0$,
  - $S 0 + z \times (S (S x) + 0)$,
  - $S y \times S 0 + S x \times 0$.
- Take as rules the following:
  \[
  \begin{align*}
  x + 0 & \rightarrow \text{Rule } x, \\
  x + S y & \rightarrow \text{Rule } S (x + y), \\
  x \times 0 & \rightarrow \text{Rule } 0, \\
  x \times S y & \rightarrow \text{Rule } x \times y + x.
  \end{align*}
  \]

(d) Term Rewriting Systems

- Term rewriting systems are special cases of reduction systems.
- They are reduction systems, which are generated by a (in many cases finite) set of rules (i.e. basic reductions).

Example Reductions
Example of a Term Rewriting System

(The system will be in fact strongly normalising and confluent).

Term Rewriting Systems

\[ \begin{align*}
  x + 0 & \rightarrow \text{Rule} \ x, \\
  x + S \ y & \rightarrow \text{Rule} \ S (x + y), \\
  x \ast 0 & \rightarrow \text{Rule} \ 0, \\
  x \ast S \ y & \rightarrow \text{Rule} \ x \ast y + x.
\end{align*} \]

The reduction relation generated by these rules allows to replace in a term
- any subterm of the form \( s + 0 \) by \( s \),
- any subterm of the form \( s + S \ t \) by \( S (s + t) \),
- any subterm of the form \( s \ast 0 \) by \( 0 \),
- any subterm of the form \( s \ast S \ t \) by \( s \ast t + s \).

So we have for instance the following reductions:
- \( 0 + S (S 0) \rightarrow S (0 + S 0) \),
- Reduce \( 0 + S \ s \) to \( S (0 + s) \) using \( s \equiv S 0 \).
- \( S (\theta + S \theta) \rightarrow S (S (\theta + \theta)) \),
- Reduce \( s + S \ t \) to \( S (s + t) \), using \( s \equiv t \equiv 0 \),
- \( S (S (\theta + \theta)) \rightarrow S (S \theta) \).

Definition of Term Rewriting Systems

A term rewriting system consists of
- a set of terms \( T \) built from variables, constants and some function symbols,
- a relation \( \rightarrow \) between terms
  (if \( r \rightarrow \) we say hat \( r \rightarrow \) is a rule),
  s.t., if \( s \rightarrow t \), then
  \( s \) is not a variable, and
  all variables in \( t \) occur in \( s \).
- The variable conditions is needed so that the theory of term rewriting systems goes through smoothly.
  This is not important for this lecture, and therefore the explanation will be omitted.
  Jump over explanation.
Condition on Variables

In the previous definition we demanded two variable conditions for \( s \rightarrow \text{Rule} \ t \):

- \( s \) is not a variable.
- If we allowed \( s \) to be a variable say \( x \), then the rule would have the form \( x \rightarrow t \).
  That would mean that any term \( r \) has a reduction, namely to \( t[x := r] \).
- All variables in \( t \) occur in \( s \).
- Assume \( y \) were a variable in \( t \) but not in \( s \).
  If we substitute in \( s \) and \( t \) all variables from \( s \) by closed terms, and obtain \( s' \) and \( t' \) then we would have that \( s' \) would have potentially infinitely many reductions, namely for any substitution of the other variables of \( t' \) by closed terms.

Condition on Variables

The second variable condition has something to do with determinism:

- Assume we have chosen a rule \( r \rightarrow s \) and chosen a substitution of variables in \( r \), which matches a term \( t \).
- Then the reduct with respect to this rule is uniquely determined.
- There are no other free variables in \( s \) which allow additional choices for substitutions.

Reduction generated by \( \rightarrow \text{Rule} \)

If we have a term rewriting system \( (T, \rightarrow \text{Rule}) \) we obtain a reduction relation \( \rightarrow \) on \( T \) as follows:

- First we construct a relation \( \rightarrow' \) obtained from reductions rules \( r \rightarrow \text{Rule} \ r' \) by substituting the variables in both \( r \) and \( r' \) by some terms.
- So the same substitutions are carried out in both \( r \) and \( r' \).
- If \( s \rightarrow' s' \) is obtained by carrying out such a substitution in \( r \rightarrow \text{Rule} \ r' \), then \( s \rightarrow' s' \) is called an instance of rule \( r \rightarrow \text{Rule} \ r' \).
Example (Instance of a Rule)

\[
\begin{align*}
x + 0 & \longrightarrow \text{Rule } x , \\
x + S y & \longrightarrow \text{Rule } S (x + y) , \\
x \ast 0 & \longrightarrow \text{Rule } 0 , \\
x \ast S y & \longrightarrow \text{Rule } x \ast y + x .
\end{align*}
\]

- \(0 + 0 \longrightarrow' 0\) is an instance, obtained by substituting in \(x + 0 \longrightarrow \text{Rule } x\) the variable \(x\) by 0.
- \(S 0 \ast S 0 \longrightarrow' S 0 \ast 0 + S 0\) is an instance, obtained by substituting in \(x \ast S y \longrightarrow \text{Rule } x \ast y + x\) the variable \(x\) by \(S 0\) and the variable \(y\) by 0.

Reduction generated by \(\longrightarrow\) Rule

Then \(s \longrightarrow s'\), if there exists an instance \(t \longrightarrow' t'\) of a rule s.t. \(s\) contains subterm \(t\), and \(s'\) is the result of substituting in \(s\) the term \(t\) by \(t'\).
- “Redex” is short for reducible expression.
- Plural of redex is redexes.
- The reductions \(s \longrightarrow s'\) obtained this way are the reductions generated by the term rewriting system.

Example 1

\[
\begin{align*}
x + 0 & \longrightarrow \text{Rule } x , \\
x + S y & \longrightarrow \text{Rule } S (x + y) , \\
x \ast 0 & \longrightarrow \text{Rule } 0 , \\
x \ast S y & \longrightarrow \text{Rule } x \ast y + x .
\end{align*}
\]

- \(0 + S (S 0) \longrightarrow S (0 + S 0)\) is obtained as follows:
  - The rule used is \(x + S y \longrightarrow \text{Rule } S (x + y)\).
  - By substituting \(x\) by 0 and \(y\) by \(S 0\) we obtain the instance \(0 + S (S 0) \longrightarrow' S (0 + S 0)\).
  - In this example, the redex is the full term \(0 + S (S 0)\) which is then reduced.

Example 2

\[
\begin{align*}
x + 0 & \longrightarrow \text{Rule } x , \\
x + S y & \longrightarrow \text{Rule } S (x + y) , \\
x \ast 0 & \longrightarrow \text{Rule } 0 , \\
x \ast S y & \longrightarrow \text{Rule } x \ast y + x .
\end{align*}
\]

- \(S (0 + S 0) \longrightarrow S (S (0 + 0))\) is obtained as follows:
  - The rule used is \(x + S y \longrightarrow \text{Rule } S (x + y)\).
  - By substituting \(x\) and \(y\) by 0 we obtain the instance \(0 + S 0 \longrightarrow' S (0 + 0)\).
  - The left hand side of our reduction \(S (0 + S 0)\) contains now the redex \(0 + S 0\).
  - By substituting it by \(S (0 + 0)\) we obtain the right hand side of the reduction, \(S (S (0 + 0))\).
Example 3

\[
\begin{align*}
x + 0 & \quad \rightarrow \text{Rule } x , \\
x + S y & \quad \rightarrow \text{Rule } S (x + y) , \\
x \ast 0 & \quad \rightarrow \text{Rule } 0 , \\
x \ast S y & \quad \rightarrow \text{Rule } x \ast y + x .
\end{align*}
\]

\[S (S (0 + 0)) \rightarrow S (S 0).\]

- The rule used is \(x + 0 \rightarrow \text{Rule } x .\)
- By substituting 0 for \(x\), we obtain the instance
  \[0 + 0 \rightarrow 0 .\]

- The left hand side of the reduction \(S (S (0 + 0))\) contains the redex \(0 + 0\).
- By substituting it by 0 we obtain the right hand side of the reduction \(S (S 0).\)