CALL-BY-VALUE SOLVABILITY

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Abstract. The notion of solvability in the call-by-value $\lambda$-calculus is defined and completely characterized, both from an operational and a logical point of view. The operational characterization is given through a reduction machine, performing the classical $\beta$-reduction, according to an innermost strategy. In fact, it turns out that the call-by-value reduction rule is too weak for capturing the solvability property of terms. The logical characterization is given through an intersection type assignment system, assigning types of a given shape to all and only the call-by-value solvable terms.

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1. INTRODUCTION

The call-by-value $\lambda$-calculus ($\lambda v$-calculus) is a paradigmatic language which captures two features present in many real functional programming languages: the call-by-value parameter passing and the lazy evaluation. The parameters are passed in a call-by-value way, when they are evaluated before being passed and a function is evaluated in a lazy way when its body is evaluated only when parameters are supplied. The real programming languages are all lazy, and almost all call-by-value (\textit{e.g.} ML [9], Scheme [13], while Haskell [14] is one of the few examples of a language using the call-by-name evaluation). Note that the call-by-value parameter passing cannot be modelled in the classical $\lambda$-calculus, since the $\beta$-reduction rule is intrinsically a call-by-name rule. The $\lambda v$-calculus is a restriction of the classical $\lambda$-calculus based on the notion of value. Values are either variables or abstractions and they represent the already evaluated terms. Since the evaluation is lazy, an abstraction is always a value, independently from

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its body. The call-by-value evaluation mechanism in the $\lambda\beta_v$-calculus is realized by defining a suitable reduction rule (the $\beta_v$-rule), which is a restriction of the classical $\beta$-rule, in the sense that $(\lambda x.M)N$ reduces to $M[N/x]$ if and only if $N$ is a value, i.e., it has been already evaluated.

The $\lambda\beta_v$-calculus and the machine for its evaluation, that we call secd, has been introduced by Plotkin [11] inspired by the seminal work of Landin [8] on the language ISWIM and the SECD machine.

In this paper we are dealing with the pure (i.e. without constants) version of the $\lambda\beta_v$-calculus. So a closed term is said valuable if its evaluation, through the secd machine, stops.

The notion of terminating programs and so of valuable terms is central for studying the operational equivalence between terms induced by the secd machine. Let a context $C[\ ]$ be a term with some occurrences of an hole, and let $C[M]$ be the term obtained from it once the holes have been filled by the term $M$. The operational equivalence is defined as follows:

$$M \simeq_v N \text{ if and only if } \forall C[\ ] \text{ such that } C[M] \text{ and } C[N] \text{ are closed, }$$

the secd machine stops on $C[M]$ if and only if it stops on $C[N]$.

This equivalence corresponds to the Liebnitz equality on programs. In fact a context $C[\ ]$ can be viewed as a partially specified program, and $C[M]$ as a program using $M$ as subprogram. So two terms are equivalent if and only if they can be replaced each other in the same program without changing its observational behaviour. In a language (like the $\lambda\beta_v$-calculus) without constants, the natural behaviour to be observed is the termination property.

Plotkin proved that the $\lambda\beta_v$-calculus enjoys some of the good properties we expected from a calculus, namely the Church-Rosser and the standardization property. But the notion of solvability, in the call-by-value setting, has never been explored. In this paper we want to study such a notion.

The notion of solvability has been introduced in the classical $\lambda$-calculus for characterizing terms with good operational behaviour. Using a programming paradigm, $M$ is solvable if and only if, for every output value $P$, there is a program $C_P[M]$, using effectively $M$ as subprogram, such that $C_P[M]$ evaluates to $P$. The fact that $C_P[M]$ uses effectively $M$ can be formalized as: not for all $Q$, $C_P[Q]$ evaluates to $P$.

In the case of classical $\lambda$-calculus, it has been proved [15] that, for all term $M$, if such a context $C[\ ]$ exists, then there is also a head context, i.e., a context of the shape:

$$(\lambda x_1 \ldots x_n. [\ ] ) M_1 \ldots M_m$$

(for some $m, n$) with the same behaviour, where $\{ x_1, \ldots, x_n \}$ is the set of free variables of $M$ ($FV(M)$). This means that the $\lambda$-terms have a functional behaviour, and so the notion of solvability can be defined in the following standard way:
A term $M$ is solvable if and only if there is a finite sequence of closed terms, $N_1, \ldots, N_m$ such that

$$(\lambda x_1 \ldots x_n. M)N_1 \ldots N_m \equiv^\beta I$$

where $FV(M) = \{x_1, \ldots, x_n\}$ and $I \equiv \lambda x.x$ is the term representing the identity function.

Solvable terms have been completely characterized from a syntactical point of view. A closed term $M$ is solvable if and only if it has a head normal form (i.e., there are integers $n, m$ and terms $M_1, \ldots, M_m$ such that $M \equiv^\beta \lambda x_1 \ldots x_n.x_i M_1 \ldots M_m$ ($1 \leq i \leq n$), for some $n$).

From an operational point of view, solvable terms are the terminating programs, in the head reduction machine [12]. From a semantic point of view, all unsolvable terms (i.e., the non terminating programs) can be all consistently equated [15]. From a logical point of view, a term $M$ is solvable if and only if it can be typed in the intersection type assignment system defined by Coppo and Dezani [3].

Let recall also the notion of solvability in the lazy $\lambda$-calculus, introduced by Abramsky and Ong [1] for modelling the call-by-name lazy evaluation. The lazy $\lambda$-calculus is the classical one, equipped with the $\beta$-reduction rule, but, in the evaluation of terms, no reduction is made under the scope of an abstraction. Abramsky and Ong in [1] noted that the notion of solvability in this setting is the same as in the call-by-name case (a term is solvable if and only if it has a head normal form). But in this case the set of solvable terms does not coincide anymore with the set of terminating terms, with respect to the lazy evaluation. Indeed the term $\lambda x. \Delta \Delta$, where $\Delta \equiv \lambda x.xx$, is unsolvable, but the lazy evaluation stops on it. In order to clarify the relation between solvable terms and termination in the lazy setting, let us recall the notion of unsolvable of order $n$ ($n \geq 0$).

Let $P$ be unsolvable. $P$ is of order 0 if and only if there is no $Q$ such that $P \equiv^\beta \lambda x.Q$; $P$ is of order $n$ if and only if $n$ is the maximum integer such that $P \equiv^\beta \lambda x_1 \ldots x_n.Q$; $P$ is of infinite order if such a $n$ does not exist. So the terminating terms in the lazy $\lambda$-calculus are the solvable terms plus the unsolvable ones of order greater than 0.

Semantically the unsolvable terms of order 0 (i.e., the non terminating programs) can be consistently equated, but a model equating all unsolvable terms is not correct with respect to the lazy operational semantics.

As far as the logical characterization of lazy solvability is concerned, it is easy to show that the logical system defined in [1] can give such a characterization.

Now let us consider the call-by-value $\lambda$-calculus.

First of all we must ask ourselves how the general notion of solvability can be specialized in this setting. In [4] (Th. 33) it has been proved that the $\lambda \beta \nu$-calculus has a functional behaviour, as the classical $\lambda$-calculus. More precisely, the operational behaviour of a term $M$ can be studied by considering just the (call-by-value) head-contexts, i.e., contexts of the shape:

$$(\lambda x_1 \ldots x_n. [ ] )P_1 \ldots P_m$$
(for some $m, n$) where all $P_i (1 \leq i \leq m)$ are values and $FV(M) = \{x_1, ..., x_n\}$. So we can define:

a term $M$ is call-by-value solvable ($v$-solvable) if and only if there is a finite sequence of closed values $N_1, ..., N_m$ such that

$$(\lambda x_1 \ldots x_n. M)N_1 \ldots N_m =_v I$$

where $FV(M) = \{x_1, ..., x_n\}$.

In this paper we will give a complete characterization, from both an operational and a logical point of view, of $v$-solvable terms.

A key observation is that, in order to characterize the class of $v$-solvable terms from an operational point of view, the $v$-reduction is too weak. In fact there are $v$-normal forms which are $v$-unsolvable, as for example the term:

$x: (y: (x:)) (x:)$

which is operationally equivalent to $x:$. So, in order to characterize operationally the $v$-solvability, a more refined tool must be designed. To do so, we extend the notion of valuability (i.e., termination) to open terms, by defining a term $M$ being potentially valuable if and only if there is a substitution $s$, replacing variables by closed values, such that $s(M)$ is valuable. It turns out that the class of the $v$-solvable terms is properly contained in that one of the potentially valuable terms. We will show that the potentially valuable terms are completely characterized through an evaluation machine, that we call inner machine, performing the classical $\beta$-reduction according to the innermost-lazy strategy. It is important to notice that the operational equivalence induced by the inner machine coincides with $\approx_s$. Another evaluation machine, the ahead machine, which is based on the previous one, gives the desired characterization of $v$-solvable terms. It turns out that a term $M$ is $v$-solvable if and only if it reduces, using the classical $\beta$-reduction with the leftmost-innermost strategy, to a term of the shape:

$$\lambda x. (\lambda y. \Delta)(xI)\Delta$$

which is operationally equivalent to $\lambda x. \Delta\Delta$. So, in order to characterize operationally the $v$-solvability, a more refined tool must be designed. To do so, we extend the notion of valuability (i.e., termination) to open terms, by defining a term $M$ being potentially valuable if and only if there is a substitution $s$, replacing variables by closed values, such that $s(M)$ is valuable. It turns out that the class of the $v$-solvable terms is properly contained in that one of the potentially valuable terms. We will show that the potentially valuable terms are completely characterized through an evaluation machine, that we call inner machine, performing the classical $\beta$-reduction according to the innermost-lazy strategy. It is important to notice that the operational equivalence induced by the inner machine coincides with $\approx_s$. Another evaluation machine, the ahead machine, which is based on the previous one, gives the desired characterization of $v$-solvable terms. It turns out that a term $M$ is $v$-solvable if and only if it reduces, using the classical $\beta$-reduction with the leftmost-innermost strategy, to a term of the shape:

$$\lambda x_1 \ldots x_n. x_1 P_1 \ldots P_m$$

where $P_i$ is potentially valuables ($1 \leq i \leq m$). Note that this definition cannot be expressed through the $\beta_v$-reduction. A preliminary version of these machines has been presented in [10].

Moreover we characterize both the potential valuability and the $v$-solvability from a logical point of view, defining an intersection type assignment system, which gives type exactly to the potentially valuable terms, and gives a type of a particular shape exactly to the $v$-solvable terms. Such a type assignment system is inspired to that one defined by [4] for reasoning about canonical denotational semantics of $\lambda\beta_v$-calculus.

Let recall that a $\lambda\beta$-theory is called sensible if it equates all unsolvable terms, and semi-sensible if it never equates a solvable term to an unsolvable one. We can extend in an obvious way this definition to a $\lambda\beta_v$-theory, calling it $v$-sensible
if it equates all the $v$-unsolvable terms, and $v$-semi-sensible if it never equates a $v$-solvable term to a $v$-unsolvable one. According to the previous definition of $v$-solvability, the $\text{secd}$-operational theory, i.e., the theory $T_{\text{secd}} = \{(M,N) | M \approx_v N\}$ is not $v$-sensible, as expected. Indeed $\Delta \Delta$ and $\lambda x \Delta \Delta$ are two different unsolvable terms which are not equated in $T_{\text{secd}}$. This depends on the fact that the $\text{secd}$ machine evaluates in a lazy way: indeed also the operational semantics of the lazy $\lambda$-calculus is not $v$-sensible. Moreover, $T_{\text{secd}}$ is not $v$-semi-sensible. In fact it turns out that it equates the identity combinator $I$ to a $v$-unsolvable term. This equivalence is not surprising, since it is a consequence of the fact that, in the minimal canonical model of $\lambda \beta \epsilon$-calculus, showed in [4], which is built by an inverse limit construction, all projections are $\lambda$-representable. We will give here a purely syntactic proof of it.

The paper is organized as follows. In Section 2 the $\lambda \beta \epsilon$-calculus and its operational semantics are recalled. In Section 3 the notions of potentially valuable and $v$-solvable term are introduced. The operational characterizations of potentially valuable and $v$-solvable terms are given in Sections 4 and 5 respectively. Section 6 contains the logical characterization. The two appendices contain the more technical proofs.

\section{The Call-by-Value $\lambda$-Calculus}

In this section we briefly recall the syntax and the operational semantics of the $\lambda \beta \epsilon$-calculus, as stated by Plotkin [11]. The $\lambda \beta \epsilon$-calculus is a restriction of the classical $\lambda$-calculus, based on the notion of value. In particular, the restriction concerns the evaluation rule, the $\beta$-rule, which is replaced by the $\beta \epsilon$-rule.

\begin{definition}
Let $\text{Var}$ be a denumerable set of variables, ranged over by $x, y, z, \ldots$
Let $\Lambda$ be the set of $\lambda$-terms, built out by the following grammar:

\[ M ::= x | MM | \lambda x.M. \]

Terms will be ranged over by $M, N, P, Q, \ldots$ A term of the form $MN$ is called application while a term of the form $\lambda x.M$ is called abstraction.

The set of values is the set $\text{Val} \subseteq \Lambda$ defined as follows:

\[ \text{Val} = \text{Var} \cup \{\lambda x.M | x \in \text{Var} \text{ and } M \in \Lambda\}. \]

It is straightforward to check that every term is of the shape:

\[ \lambda x_1 \ldots x_n . \zeta M_1 \ldots M_m \]

for some $n, m \geq 0$, where $M_i \in \Lambda$, and $\zeta$ is either a variable or an abstraction.

\begin{notation}
Free and bound variables are defined as usual, $\text{FV}(M)$ denotes the set of free variables of the term $M$ and $\Lambda^0 \subseteq \Lambda$ denotes the set of closed terms,
\end{notation}
i.e., terms whose set of free variables is empty. Moreover $\text{Val}^0 \subset \text{Val}$ denotes the subset $\text{Val} \cap \Lambda^0$. A context (denoted by $C[\ ]$) is a term with some occurrences of a hole; it can be built by a grammar obtained from that one for $\lambda$-terms by adjoining the hole to the set of variables. $C[M]$ denotes the context $C[\ ]$ once every occurrence of the hole has been replaced by the term $M$. Note that in the replacement free variables can be captured, and so they can become bound. As usual, terms are considered modulo $\alpha$-conversion, i.e., modulo renaming of bound variables. Moreover $\lambda x_1 \ldots x_n . M$ is an abbreviation for $\lambda x_1 . (\lambda x_2 . (\ldots (\lambda x_n . M)))$ and $M_1 \ldots M_m$ is an abbreviation for $((\ldots ((M_1 M_2) M_3) \ldots) M_m)$. $\equiv$ denotes the syntactical identity on terms.

**Definition 2.3.** The call-by-value evaluation rule is defined as follows:

\[(\beta_v) \quad (\lambda x . M)N \rightarrow M[N/x] \quad \text{if} \quad N \in \text{Val}\]

where $M[N/x]$ denotes the simultaneous replacement of every free occurrences of $x$ in $M$ by $N$, renaming bound variables in $M$ to avoid variable clash.

Let $\rightarrow_v$, $\rightarrow^*_v$ and $=_v$ denote respectively the contextual closure of the $\beta_v$-rule, the reflexive and transitive closure of $\rightarrow_v$, and the reflexive, symmetric and transitive closure of $\rightarrow_v$.

The $\beta_v$ reduction satisfies both the Church-Rosser property and the Standardization property (see [11]).

The evaluation of a program (closed term) is formalized through a reduction machine, which we call seed machine for pointing out that it is equivalent (w.r.t. the termination property) to the S.E.C.D. machine defined by Landin for evaluating expressions [8], once its input is restricted to pure $\lambda$-calculus terms. We give here a logical presentation of this machine, i.e., the machine is defined as a set of logical rules, and the evaluation process is mimicked by a logical derivation.

The operational equivalence between terms is determined by observing the termination of computations carried out by the seed machine.

**Definition 2.4.** i) The seed-machine is a set of rules proving statement of the shape:

\[ M \Downarrow_s N \]

where $M \in \Lambda^0$ and $N \in \text{Val}$. The rules are:

\[
\begin{align*}
\lambda x . M \Downarrow_s \lambda x . M & \quad \text{(abs)} \\
M \Downarrow_s \lambda x . M' & \quad N \Downarrow_s N' & \quad M'[N'/x] \Downarrow_s P & \quad \text{(app)}
\end{align*}
\]

If $M \Downarrow_s N$, we will say that $M$ is the *input* of the seed machine and $N$ is the corresponding *output*.

Let $M \Downarrow_s$ be an abbreviation for: $\exists N$ such that $M \Downarrow_s N$. If $M \Downarrow_s$ we will say that $M$ is *valuable*. 


ii) Two terms \( M \) and \( N \) are \textit{secd-operationally equivalent} (\( M \approx_{s} N \)) if and only if for all context \( C[ \ ] \) such that \( C[M], C[N] \in \Lambda^0 \),

\[
C[M] \downarrow_s \iff C[N] \downarrow_s.
\]

It is immediate to verify that the secd-machine is deterministic, \textit{i.e.}, if \( M \downarrow_s \) then there is exactly one \( N \) such that \( M \approx_{s} N \) and moreover there is exactly one derivation proving \( M \downarrow_s N \). So, if \( M \downarrow_s \), then we can define the \textit{number of steps} of the secd-machine when filled with input \( M \) (notation: \( \text{steps}_s(M) \)) as the number of applications of rules in the derivation proving \( M \downarrow_s N \).

It can be checked that the secd-machine reduces at the every step the leftmost outermost \( \beta_v \)-redex occurring in the input term not inside the scope of an abstraction, until a value is reached. The following proposition holds:

\textbf{Proposition 2.5.} \( M \in \Lambda^0 \), \( M \rightarrow^*_{v} N \) and \( N \in \text{Val} \) if and only if \( M \) is valuable.

\textit{Proof.} By the standardization property of \( \beta_v \)-reduction, see [11]. \hfill \Box

3. Potentially valuable and \( v \)-solvable terms

In this section, both the notions of potentially valuable and \( v \)-solvable term are introduced, and their relation is discussed. The notion of potentially valuable term is the extension to open terms of the notion of termination in the secd machine. Note that this extension cannot be defined in the standard way, by defining an open term being potentially valuable if its closure is valuable, since the secd machine evaluates terms in a lazy way, so all abstractions are terminating.

\textbf{Definition 3.1.} A term \( M \) is \textit{potentially valuable} if and only if there is a substitution \( s \), replacing variables by closed values, such that \( s(M) \) is valuable.

It is immediate to verify that a closed term is potentially valuable if and only if it is valuable.

Now, let us define the notion of \( v \)-solvability, for grasping the functional behaviour of terms.

\textbf{Definition 3.2.} A term \( M \) is \textit{\( v \)-solvable} if and only if there are values \( N_1, \ldots, N_n \in \text{Val}^0 \) such that:

\[
(\lambda x_1 \ldots x_m.M)N_1 \ldots N_n =_v I
\]

where \( I \equiv \lambda x.x \) and \( \text{FV}(M) = \{x_1, \ldots, x_m\} \).

A term is \textit{\( v \)-unsolvable} if and only if it is not \( v \)-solvable.

\textbf{Lemma 3.3.} The class of \( v \)-solvable terms is properly included in the class of potentially valuable terms.

\textit{Proof.} Let first prove the inclusion. Let \( M \) be \( v \)-solvable, so for some closed values \( N_1, \ldots, N_n \), \( (\lambda x_1 \ldots x_m.M)N_1 \ldots N_n =_v I \). Without loss of generality, we can assume \( m \leq n \). In fact, if \( m > n \), then there are \( P_j \) \((n+1 \leq j \leq m)\) such that

\[
(\lambda x_1 \ldots x_m.M)N_1 \ldots N_n P_{n+1} \ldots P_m =_v I;
\]
just take \( P_j = I \), for all \( j \). Since \( I \) is a normal form, \( (\lambda x_1 \ldots x_m. M) N_1 \ldots N_n \rightarrow^* I \) and so \( M[N_1/x_1, \ldots, N_m/x_m] N_{m+1} \ldots N_n \rightarrow^* I \). By Proposition 2.5,

\[
M[N_1/x_1, \ldots, N_m/x_m] N_{m+1} \ldots N_n \Downarrow
\]

and this implies \( M[N_1/x_1, \ldots, N_m/x_m] \Downarrow \). The inclusion is proper, since \( \lambda x. \Delta \Delta \) is valuable, and so potentially valuable, but clearly \( v \)-unsolvable.

### 4. Operational characterization of potentially valuable terms

In this section a new reduction machine, the inner-machine, is introduced, which operationally characterizes the potentially valuable terms, in the sense that it stops if and only if the input term is potentially valuable. The shape of the output results of such a machine, which we call canonical terms, is particularly interesting.

**Definition 4.1.**

i) A term \( M \) is canonical if and only if it is either a value or of the shape:

\[
xM_1 \ldots M_m \quad (m \geq 0)
\]

where \( M_i (1 \leq i \leq m) \) is canonical. Let \( C \) be the set of canonical terms.

ii) The inner-machine is a set of rules proving statements of the shape:

\[
M \Downarrow_i N
\]

where \( M \in \Lambda \) and \( N \in C \). The rules are:

\[
m \geq 0 \quad M_j \Downarrow_i N_j (1 \leq j \leq m) \quad \text{(var)}
\]

\[
Q \Downarrow_i R \quad P[R/z]M_1 \ldots M_m \Downarrow_i N \quad (\text{head})
\]

\[
\lambda x. Q \Downarrow_i \lambda x. Q \quad \text{(lazy)}.
\]

Let \( M \Downarrow_i \) be an abbreviation for: \( \exists N \text{ such that } M \Downarrow_i N \).

If \( M \Downarrow_i N \), then \( M \) is the input of the \( i \)-machine and \( N \) is the corresponding output.

It is easy to prove that the inner-machine is well-defined, i.e., if \( M \Downarrow_i N \) then \( N \in C \), and moreover the machine is deterministic. So the notion of the number of steps (defined for the seed-machine in the previous section) can be extended to the \( i \)-machine in a straightforward way: if \( M \Downarrow_i \), \( \text{step}_i(M) \) denotes the number of steps performed by the \( i \)-machine on input \( M \).
Note that the inner-machine executes the classical $\beta$-reduction (call by name) with an innermost-lazy strategy. In fact it performs at every step the leftmost innermost $\beta$-redex not inside the scope of a $\lambda$-abstraction, until either an abstraction or a head variable is reached, and, in the last case, it performs the same reduction strategy in parallel inside all the arguments.

Let us introduce a new reduction rule:

\[(\lambda z.M)N \rightarrow_{\text{inner}}^* M[N/z] \quad \text{if} \quad N \in C.\]

Let $\rightarrow_\iota$, $\rightarrow_\iota^*$ and $\equiv_\iota$ denote respectively the closure under application of the $\rightarrow_\iota$, the reflexive and transitive closure of $\rightarrow_\iota$ and the reflexive, symmetric and transitive closure of $\rightarrow_\iota$.

Note that $\rightarrow_\iota$, $\rightarrow_\iota^*$ and $\equiv_\iota$ are not contextual closed, but they are just closed under application. Indeed, the reduction relation obtained by the contextual closure of $\rightarrow_\iota$ is not Church-Rosser (e.g., the term $(\lambda x.(\lambda z.I)(x\Delta))\Delta$ would reduce both to $I$ and $(\lambda z.I)(\Delta\Delta)$, which do not have a common reduct).

$\rightarrow_\iota^*$, as have been defined, is Church-Rosser (it can be easily proved), and moreover, being not closed under abstraction, it is intrinsically lazy. As far as the example before is concerned, note that the term $(\lambda x.(\lambda z.I)(x\Delta))\Delta$ has just one $i$-redex, and it $i$-reduces only to $(\lambda z.I)(\Delta\Delta)$.

The inner-machine can be alternatively described as performing the $\rightarrow_\iota$ reduction. More precisely, it performs the leftmost outermost $i$-redex not inside the scope of a $\lambda$-abstraction, until either an abstraction or a head variable is reached, and in this last case it performs the same reduction strategy inside all the arguments. Moreover canonical terms are lazy normal forms with respect to the $i$-reduction rule, i.e., a canonical term does not contain $i$-redexes, but inside the scope of a $\lambda$-abstraction. The following proposition clarifies the relation between the $i$-machine and the $i$-reduction.

Proposition 4.2. $M \downarrow_i N$ if and only if $M \rightarrow_\iota^* N$ and $N$ is canonical.

Proof. ($\Rightarrow$) By induction on the definition of the $i$-machine. ($\Leftarrow$) By induction on the length of the reduction $M \rightarrow_\iota^* N$.

The behaviour of the inner-machine and of the seed-machine coincide on closed terms, as proved by the following proposition.

Proposition 4.3. Let $M \in \Lambda^0$. Then $M \downarrow_i N$ if and only if $M \downarrow_s N$.

Proof. ($\Rightarrow$) By induction on $\text{step}_i(M)$. In case the last used rule is (lazy) then the proof is obvious. The last used rule cannot be (var), since $M$ is closed. Let the last used rule be (head), and let the derivation be:

\[
\frac{Q \downarrow_i R \quad P[R/x]M_1 \ldots M_m \downarrow_i N}{(\lambda x.P)QM_1 \ldots M_m \downarrow_i N} \quad \text{(head)}.
\]

By induction $Q \downarrow_s R$ and $P[R/x]M_1 \ldots M_m \downarrow_s N$. This implies there is a derivation:
\[
\frac{P[R/x]M_1\ldots M_{m-1} \Downarrow_{\alpha} \lambda z. S \ M_m \Downarrow_{\alpha} T \ S[T/z] \Downarrow_{\alpha} N}{P[R/x]M_1\ldots M_m \Downarrow_{\alpha} N}\text{(app)}.
\]

We will prove that \(P[R/x]M_1\ldots M_m \Downarrow_{\alpha} N\) implies \((\lambda x. P)QM_1\ldots M_m \Downarrow_{\alpha} N\), by induction on \(m\). The first step is obvious, by the rule (app). For the induction step, looking at the derivation showed before, we can assume \((\lambda x. P)QM_1\ldots M_{m-1} \Downarrow_{\alpha} \lambda z. S\). So we can build the derivation:

\[
\frac{(\lambda x. P)QM_1\ldots M_m \Downarrow_{\alpha} \lambda z. S \ M_m \Downarrow_{\alpha} T \ S[T/z] \Downarrow_{\alpha} N}{(\lambda x. P)QM_1\ldots M_m \Downarrow_{\alpha} N}\text{(app)}.
\]

(\(\Leftarrow\)) By induction on \text{step}_{\alpha}(M). If the last used rule is (\text{abs}) then the proof is obvious. Otherwise, let the derivation be:

\[
\frac{P \Downarrow_{\alpha} \lambda x.R \ Q \Downarrow_{\alpha} S \ R[S/x] \Downarrow_{\alpha} N}{PQ \Downarrow_{\alpha} N}\text{(app)}.
\]

By induction \(P \Downarrow_{\alpha} \lambda x.R, Q \Downarrow_{\alpha} S\) and \(R[S/x] \Downarrow_{\alpha} N\). So \((\lambda x. R)Q \Downarrow_{\alpha} N\) (by the rule (\text{head})). We will prove, by induction on \text{step}_{\alpha}(P), that \((\lambda x. R)Q \Downarrow_{\alpha} N\) and \(P \Downarrow_{\alpha} \lambda x.R\) imply \(PQ \Downarrow_{\alpha} N\). The case \text{step}_{\alpha}(P) = 1 is obvious. Otherwise, let \(P \equiv (\lambda z. S)TP_1\ldots P_r (r \geq 0)\). Then:

\[
\frac{T \Downarrow_{\alpha} T' \ S[T'/z]P_1\ldots P_r \Downarrow_{\alpha} \lambda x.R}{(\lambda x. S)TP_1\ldots P_r \Downarrow_{\alpha} \lambda x.R}\text{(head)}.
\]

By induction, \(S[T'/z]P_1\ldots P_r \Downarrow_{\alpha} \lambda x.R\) and \((\lambda x. R)Q \Downarrow_{\alpha} N\) imply \(S[T'/z]P_1\ldots P_r Q \Downarrow_{\alpha} \lambda x. R\). So we can build the following derivation:

\[
\frac{T \Downarrow_{\alpha} T' \ S[T'/z]P_1\ldots P_r Q \Downarrow_{\alpha} \lambda x.R}{(\lambda x. S)TP_1\ldots P_r Q \Downarrow_{\alpha} \lambda x. R}\text{(head)}.
\]

In order to prove that the inner-machine completely characterizes the potentially valuable terms, we need some lemmas. Moreover, for proving them, we need to introduce a measure to be used for currying out the induction. Informally such a measure, that we call \text{weight}, is an upper bound to both the number of lazy \(\beta\)-reductions and of \(i\)-reductions needed for reducing a term to a value, if it is possible.

**Definition 4.4.** The weight of \(M\) (denoted by \(\langle M \rangle\)), is the partial function defined as follows:

- \((M) = 0\) if \(M \in Val\)
- \(((\lambda x.M_0)M_1\ldots M_m) = 1 + (M_1) + (M_0[M_1/x]M_2\ldots M_m)\).

**Proposition 4.5.**

i) \(M \rightarrow^* \alpha N \in Val\) implies \(\langle M \rangle \geq \langle N \rangle\).

ii) \(M \rightarrow^*_\beta N \in Val\) implies \(\langle M \rangle \geq \langle N \rangle\).

iii) \(M \rightarrow^*_i N \in Val\) implies \(\langle M \rangle \geq \langle N \rangle\).
Proof. See Appendix A. □

**Lemma 4.6.** Let $M \equiv (\lambda x.M_0)M_1\ldots M_m \in \Lambda^0$. If there is $N \in \text{Val}$ such that $M \rightarrow^* N$ then, for all $i$ $(1 \leq i \leq m)$, $M_i \rightarrow^*_v N_i$, for some value $N_i$, and $\langle M_i \rangle < \langle M \rangle$.

**Proof.** By Proposition 4.3, $M \rightarrow^*_v N \in \text{Val}$ implies $M \Downarrow_i N$. So we will give the proof by induction on $\text{step}_i(M)$. If $\text{step}_i(M) = 1$, then $m = 0$, and the proposition is vacuously true. Otherwise, the derivation of $M \Downarrow_i N$ is of the shape:

$$\frac{M_1 \Downarrow_i R_1\text{M}_0[R_1/x]\text{M}_2\ldots \text{M}_m \Downarrow_i N}{(\lambda x.\text{M}_0)\text{M}_1\ldots \text{M}_m \Downarrow_i N} \quad \text{(head)}.$$

By Proposition 4.3, both $M_1$ and $M_0[R_1/x]M_2\ldots M_m$ are reducible to a value, and $\langle M_i \rangle < \langle M \rangle$ follows by definition of $\langle M \rangle$, while $\langle M_i \rangle < \langle M \rangle$ $(2 \leq i \leq m)$ follows by induction.

**Lemma 4.7.** Let $M \in \Lambda$ and let $\text{FV}(M) \subseteq \{x_1,\ldots, x_n\}$. If there are $P_1,\ldots, P_n \in \text{Val}^0$ and $M \in \text{Val}^0$ such that $M[P_1/x_1,\ldots, P_n/x_n] \rightarrow^*_v M'$, then there is $N \in \Lambda$ such that both $M \Downarrow_i N$ and $N[P_1/x_1,\ldots, P_n/x_n] \rightarrow^*_v M'$.

**Proof.** In this proof we will denote $R[P_1/x_1,\ldots, P_n/x_n]$ by $R'$, for every $R \in \Lambda$.

The proof is carried out by induction on the weight $\langle M' \rangle = k$. Note that $M' \in \Lambda^0$.

$k = 0$: Then $M'$ is already a value and, since it is closed, it must be $M' \equiv \lambda z.P'$. There are two cases:

1. $M \equiv x_j$ and $P_j \equiv \lambda z.P'$. So $x_j \Downarrow_i x_j$, by the (var) rule, with $m = 0$, and $x_j[y/x_j] \rightarrow^*_v y$.

2. $M \equiv \lambda z.P$. This case is obvious since the inner-machine stops by the (lazy) rule.

$k > 0$: Then $M' \equiv (\lambda z.P')M'_1\ldots M'_m$ $(m > 0)$. Then two cases are possible (with respect to the shape of $M'$):

1. $M \equiv x_jM_1\ldots M_m$ and $P_j \equiv \lambda z.P'$. So

$$(x_jM_1\ldots M_m)[P_1/x_1,\ldots, P_n/x_n] \equiv P_jM'_1\ldots M'_m \rightarrow^*_v \tilde{M}.$$

By Lemma 4.6 and Proposition 2.5, $M'_i \rightarrow^*_v \tilde{M}_i \in \text{Val}$ and $\langle M'_i \rangle < k$, so by induction there is $N_i$ such that $M_i \Downarrow_i N_i$ and $N_i[P_1/x_1,\ldots, P_n/x_n] \rightarrow^*_v \tilde{M}_i$ $(1 \leq i \leq n)$. So $x_jM_1\ldots M_m \Downarrow_i x_jN_1\ldots N_m$ by rule (var) and

$$P_jN'_1\ldots N'_m \rightarrow^*_v P_j\tilde{M}_1\ldots \tilde{M}_m \rightarrow^*_v \tilde{M}.$$

2. $M \equiv (\lambda z.P)M_1\ldots M_m$.

By Lemma 4.6 and Proposition 2.5, $M'_1 \rightarrow^*_v \tilde{M}_1 \in \text{Val}$, and $\langle M'_1 \rangle < k$. So by induction there is $R$ such that $M_1 \Downarrow_i R$ and $R' \rightarrow^*_v M_1$. 

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Lemma 4.8. Let \( P \in L \), \( FV(P) \subseteq \{x_1, \ldots, x_n\} \) and \( Q' = \lambda x_1 \ldots x_{r+1}, x_{r+2} \ldots x_{n} \)

Clearly \( P'[M_1/z]M_2 \ldots M'_m \rightarrow^*_v M \), thus \( P'[M_1/z]M_2 \ldots M'_m \rightarrow^*_v M \).

Moreover \( \langle P'[M_1/z]M_2 \ldots M'_m \rangle < k \) by definition of weight, so by

Proposition 4.5.ii \( \langle P'[R'/z]M_2 \ldots M'_m \rangle < k \), furthermore \( P'[R'/z]M_2 \ldots M'_m \rightarrow^*_v M \), because \( P'[M_1/z]M_2 \ldots M'_m \rightarrow^*_v M \) and \( R' \rightarrow^*_v M_1 \). So by induction \( P[R/z]M_2 \ldots M_m \not\vdash_i N \), for some \( N \), and

\[ N[P_1/x_1, \ldots, P_n/x_n] \rightarrow^*_v M. \]

Then \( (\lambda z.P)M_1 \ldots M_m \not\vdash_i N \), by the rule head.

\[ \square \]

**Proof.** By induction on \( \text{step}_i(M) \). The case \( \text{step}_i(M) = 1 \) is trivial.

In the case the last applied rule is \((\text{var})\) the result is obvious.

Let the last applied rule be

\[
\frac{Q \not\vdash_i R \quad P[R/z]M_1 \ldots M_m \not\vdash_i N}{(\lambda z.P)QM_1 \ldots M_m \not\vdash_i N} \hspace{1cm} \text{(head}).
\]

Let \( q = \text{step}_i(Q) \). By induction \( \forall r \geq q, Q[O_r/x_1, \ldots, O_r/x_n] \rightarrow^*_v Q \in Val^0 \), and by Lemma 4.7, \( R[O_r/x_1, \ldots, O_r/x_n] \rightarrow^*_v Q \).

Let \( p = \text{step}_i(P[R/z]M_1 \ldots M_m) \). By induction, \( \forall h \geq p: \)

\[
(P[R/z]M_1 \ldots M_m)[O_h/x_1, \ldots, O_h/x_n] \rightarrow^*_v N \in Val.
\]

In particular, since \( \text{step}_i(M) = 1 + q + p \), for all \( \rho \geq \text{step}_i(M) \), both

\[
Q' \equiv Q[O_r/x_1, \ldots, O_r/x_n] \rightarrow^*_v Q' \in Val^0
\]

and

\[
R' \equiv R[O_r/x_1, \ldots, O_r/x_n] \rightarrow^*_v Q' \in Val^0.
\]

Let \( P' \equiv P[O_r/x_1, \ldots, O_r/x_n] \) and \( M' \equiv M_1[O_r/x_1, \ldots, O_r/x_n] \);

\[
(P[R/z]M_1 \ldots M_m)[O_r/x_1, \ldots, O_r/x_n] \equiv P'[R'/z]M'_1 \ldots M'_m \rightarrow^*_v
\]

\[
\rightarrow^*_v P'[Q'/z]M'_1 \ldots M'_m \rightarrow^*_v \tilde{M},
\]

for some \( \tilde{M} \in Val^0 \). So

\[
((\lambda z.P)QM_1 \ldots M_m)[O_r/x_1, \ldots, O_r/x_n] \equiv (\lambda z.P')Q'M'_1 \ldots M'_m \rightarrow^*_v
\]
\[(\lambda z. P^0 Q^0 M_1^0 \ldots M_m^0 \rightarrow_v P^0[Q^0/z]M_1^0 \ldots M_m^0 \rightarrow^* M \in Val^0).\]

Now we are able to prove the characterization theorem.

**Theorem 4.9 (inner property).** \(M \downarrow_i \text{ if and only if } M \text{ is potentially valuable.}\)

**Proof.** (\(\Leftarrow\)) The proof follows directly from Lemma 4.8. (\(\Rightarrow\)) By definition, \(M\) potentially valuable means that there is a substitution \(s\), replacing variables by closed values, such that \(s(M)\) is valuable. By Proposition 2.5, this implies \(s(M) \rightarrow^* N \in Val^0\), and, by Lemma 4.7, this implies \(M \downarrow_i\).

The inner-machine induces an operational equivalence on terms, defined in the usual way as follows.

**Definition 4.10.** Let \(M, N \in \Lambda\). \(M\) and \(N\) are \(i\)-operationally equivalent \((M \approx_i N)\), if and only if for all context \(C[\ ]\) such that \(C[M], C[N] \in \Lambda^0\), \(C[M] \downarrow_i \Leftrightarrow C[N] \downarrow_i\).

By Proposition 4.3, \(\approx_s\) and \(\approx_i\) coincide.

It can be interesting to consider an extension of the operational equivalence, by dropping the restriction that contexts must be closing. Let define:

**Definition 4.11 (open-equivalence).** The term \(M\) and \(N\) are \(i\)-open-operationally equivalent \((M \approx^o_i N)\) if and only if for all context \(C[\ ]\), \(C[M] \downarrow_i \Leftrightarrow C[N] \downarrow_i\).

**Proposition 4.12.** \(M \approx_i N \text{ if and only if } M \approx^o_i N\).

**Proof.** Both directions can be proved by contraposition.

\((\Leftarrow\): Obvious.

\((\Rightarrow\): Let \(C[M] \downarrow_i\) and \(C[N] \not\downarrow_i\), for some \(C[\ ]\) not necessarily closing. \(C[M]\) is potentially valuable, so there is a sequence of closed values \(P_1, \ldots, P_n\) such that, if \(\text{FV}(C[M]) \subseteq \{x_1, \ldots, x_n\}\), then \(C[M][P_1/x_1, \ldots, P_n/x_n] \not\downarrow_i\). Then the closing context, separating \(M\) and \(N\) is

\[C'[\ ] \equiv (\lambda x_1 \ldots x_m. C[\ ]) P_1 \ldots P_n \frac{I}{m-n} \]

where \(\text{FV}(C[N]) \cup \text{FV}(C[M]) = \{x_1, \ldots, x_m\}\).

Note that in general the equivalence induced by closing contexts does not coincide with that one induced by all context. For example, let us consider the machine which takes a \(\lambda\)-term as input, performs at every step the leftmost outermost \(\beta_i\)-redex not inside the scope of a \(\lambda\)-abstraction and stops on the lazy \(\beta_i\)-normal form. For closed terms this machine is equivalent to the seed-machine, so it induces the same equivalence. Consider the terms \(P_0 \equiv \lambda y.(\lambda x. \Delta \Delta)(y I)\) and \(P_1 \equiv \lambda y. \Delta \Delta\).
which are \( \approx_{x} \). Let \( C[\ ] \equiv [\ ](\lambda x.x(zI)) \). Then the previous described machine stops on \( C[P_{0}] \), while does not stop on \( C[P_{1}] \).

It is important to notice that the behaviour of the inner-machine is in some sense anomalous, since \( M \downarrow_{i} N \) does not necessarily imply \( M \approx_{i} N \). A counter example is the term \((\lambda yz.I)(xI)\): it is immediate to check that \((\lambda yz.I)(xI) \downarrow_{i} \lambda z.I\), while the context \((\lambda x.[\ ])(\lambda x.\Delta\Delta)\) separates the two terms.

5. OPERATIONAL CHARACTERIZATION OF \( v \)-SOLVABILITY

In this section the operational characterization of the \( v \)-solvability is given, through a reduction machine, the ahead machine.

Such a reduction machine performs the \( \beta \)-reduction and uses the inner-machine as submachine.

Definition 5.1. i) A term \( M \) is a \( v \)-head normal form (v.h.n.f) if and only if it has the following shape:

\[
\lambda x_{1} \ldots x_{n}.xM_{1} \ldots M_{m}
\]

where \( M_{i} \in \mathcal{C} \quad (1 \leq i \leq m) \). Let \( \mathcal{VH} \) be the set of \( v \)-head normal forms.

ii) The ahead-machine is a set of rules proving statements of the shape:

\[
M \downarrow_{a}^{0} N
\]

where \( M \in \mathcal{A} \) and \( N \in \mathcal{VH} \). The rules define an auxiliary machine too, proving statements of the shape \( M \downarrow_{a}^{1} N \). The set of rules is the following, where \( k \in \{0,1\} \):

\[
\begin{align*}
\frac{m \geq 0 \quad M_{i} \downarrow_{a}^{1} N_{i} \quad (1 \leq i \leq m)}{xM_{1} \ldots M_{m} \downarrow_{a}^{k} xN_{1} \ldots N_{m} \quad (\text{var})} \\
\frac{Q \downarrow_{a}^{1} R \quad P[R/z]M_{1} \ldots M_{m} \downarrow_{a}^{k} N}{(\lambda z.P)QM_{1} \ldots M_{m} \downarrow_{a}^{k} N \quad (\text{head})} \\
\frac{\lambda x.Q \downarrow_{a}^{1} \lambda x.Q}{(\text{lazy})} \\
\frac{P \downarrow_{a}^{0} Q}{\lambda x.P \downarrow_{a}^{0} \lambda x.Q \quad (\lambda 0) .}
\end{align*}
\]

Let \( M \downarrow_{a}^{k} \) be an abbreviation for \( M \downarrow_{a}^{k} N \), for some \( N \).
It is easy to check that the definition is correct, i.e., $M \downarrow^0 N$ implies $N \in \mathcal{VH}$. Furthermore, note that the machine of level 1 is the inner-machine, i.e. $M \downarrow^0 N$ if and only if $M \downarrow_1 N$. The behaviour of the ahead-machine is not completely lazy: it enters under the external abstraction (if any) and then it works exactly as the inner-machine. In order to give a precise characterization, in terms of reductions, of the behaviour of the ahead-machine, we need to introduce a new reduction rule. Let $\rightarrow_I$, $\rightarrow^*_I$ and $=_I$ be the not lazy version of $\rightarrow$, $\rightarrow^*$ and $=$, respectively; namely $\rightarrow_I$, $\rightarrow^*_I$ and $=_I$ denote respectively the contextual closure of $\rightarrow_{inner}$, the reflexive and transitive closure of $\rightarrow_I$ and the reflexive, symmetric and transitive closure of $\rightarrow_I$.

**Proposition 5.2.** $M \downarrow^0 N$ if and only if $N$ is of the shape $\lambda x_1 \ldots x_n . xN_1 \ldots N_m$, and $M \rightarrow^*_I \lambda x_1 \ldots x_n . xM_1 \ldots M_m$, and $M_j \downarrow_1 N_j$ $(1 \leq j \leq m)$.

Note that the fact that the $\rightarrow_I$ reduction is not Church-Rosser does not create any problem, since the ahead-machine performs a particular strategy on it. So this new machine is deterministic, and if $M \downarrow^k$ then $\text{step}^k_\downarrow(M)$ is the numbers of steps of the derivation proving $M \downarrow^k N$.

For proving the desired characterization, we need two lemmas.

**Lemma 5.3.** Let $M \in \Lambda$ and $FV(M) \subseteq \{x_1, \ldots, x_n\}$. If there are $P_1, \ldots, P_k \in Val^0$ such that $(\lambda x_1 \ldots x_n . M)P_1 \ldots P_k \rightarrow^*_I$ then $M \downarrow^0 N$ and $(\lambda x_1 \ldots x_n . N)P_1 \ldots P_k \rightarrow^*_I$.

**Proof.** As showed in the proof of Lemma 3.3, it is always possible to assume $k \geq n$.

Let $S \equiv (\lambda x_1 \ldots x_n . M)P_1 \ldots P_k$ and, for every $R \in \Lambda$, let $R'$ be $R[P_1/x_1, \ldots , P_n/x_n]$. The proof is given by induction on the following pair: $(|S|, \text{number of symbols in } M)$, ordered by the lexicographical order.

(0, 0): Then $S \in Val$ and the proof is trivial by rule (λ0).

(0, 1, 1): Let analyze all possible shapes of the term $M$.

- $M \equiv x_j M_1 \ldots M_m$ $(m \geq 1)$. By hypothesis there are $P_1, \ldots , P_k \in Val^0$ such that $(\lambda x_1 \ldots x_n . x_j M_1 \ldots M_m)P_1 \ldots P_k \rightarrow^*_I P_j M'_1 \ldots M'_m P_{n+1} \ldots P_k \rightarrow^*_I I$

where, by Lemma 4.6 and Proposition 2.5, $M' \rightarrow^*_I \tilde{M}_i \in Val$. So by Lemma 4.7 (using the fact that $\downarrow^1_1$ coincides with $\downarrow_1$) we can state $M_i \downarrow^1_1 N_i$ and $N_i \equiv N_i[P_1/x_1, \ldots , P_n/x_n] \rightarrow^*_v \tilde{M}_i$.

So $x_j M_1 \ldots M_m \downarrow^0_1 x_j N_1 \ldots N_m$, by rule (var), and $(\lambda x_1 \ldots x_n . x_j N_1 \ldots N_m)P_1 \ldots P_k \rightarrow^*_v P_j N'_1 \ldots N'_m P_{n+1} \ldots P_k \rightarrow^*_v P_j \tilde{M}_1 \ldots \tilde{M}_m P_{n+1} \ldots P_k \rightarrow^*_I$.

- $M \equiv (\lambda z . P)QM_1 \ldots M_m$ $(m \geq 0)$. Then

$$S \equiv (\lambda x_1 \ldots x_n . (\lambda z . P)QM_1 \ldots M_m)P_1 \ldots P_k \rightarrow^*_I.$$
\[ \rightarrow_{v}^{*} (\lambda z.P')Q'M_1\ldots M_m P_{n+1}\ldots P_k \rightarrow_{v}^{*} (\lambda z.P')Q'M_1\ldots M_m P_{n+1}\ldots P_k \rightarrow^{*} I \]

where \( Q' \rightarrow_{v}^{*} Q \in Val \). Since \( Q' \rightarrow_{v}^{*} Q \in Val \), by Lemma 4.7 \( \Downarrow_{i} R \) and

\[ R' \rightarrow_{v}^{*} Q \in Val. \]

Let us remind that \( Q \Downarrow_{i} R \) coincides with \( Q \Downarrow_{i} R \) (\( \ast \)). Observe that

\[ (\lambda z.P')Q'M_1\ldots M_m P_{n+1}\ldots P_k \rightarrow_{\beta} P'[Q'/z]M_1\ldots M_m P_{n+1}\ldots P_k \rightarrow_{\beta}^{*} \]

\[ - \rightarrow_{\beta}^{*} P'[R'/z]M_1\ldots M_m P_{n+1}\ldots P_k \rightarrow_{v}^{*} I \quad (**) \]

Since

\[ (S) = n + \langle P_1 \rangle + \ldots + \langle P_n \rangle + \langle (\lambda z.P')Q'M_1\ldots M_m P_{n+1}\ldots P_k \rangle \]

\[ = n + \langle P_1 \rangle + \ldots + \langle P_n \rangle + 1 + \langle Q' \rangle + \langle P'[Q'/z]M_1\ldots M_m P_{n+1}\ldots P_k \rangle \]

\[ = (\text{by Prop. 4.5.ii}) \]

\[ \langle P'[Q'/z]M_1\ldots M_m P_{n+1}\ldots P_k \rangle \geq \langle P'[R'/z]M_1\ldots M_m P_{n+1}\ldots P_k \rangle \]

we can apply the induction hypothesis (**) obtaining \( P[R/z]M_1\ldots M_m \Downarrow_{a} N \) (\( \ast \ast \ast \)) and \( (\lambda x_1\ldots x_n.N)P_1\ldots P_k \rightarrow_{v}^{*} I \).

\( \ast \) and \( \ast \ast \ast \) together imply \( (\lambda x.P)QM_1\ldots M_m \Downarrow_{a} N \), and the proof is given.

\[ \ast \quad M \equiv \lambda x.P. \] This case is straightforward by induction on \( s \).

\[ \square \]

**Lemma 5.4.** Let \( M \in \Lambda \) and let \( FV(M) \subseteq \{x_1,\ldots, x_n\} \). If \( M \Downarrow_{a}^{0} \) then \( \forall r > \max\{n, \text{step}_{a}^{0}(M)\}, \exists h \geq 0 \) such that

\[ (\lambda x_1\ldots x_n.M) \underbrace{Q',\ldots, Q'}_{r} \rightarrow_{v}^{*} O^{h} \]

where \( O^{h} \equiv \lambda x_1\ldots x_{k+1}.x_{k+1}. \)

**Proof.** First of all, observe that \( P \Downarrow_{i} Q \) and \( FV(P) \subseteq \{x_1,\ldots, x_n\} \) imply (by Lem. 4.8) that \( \forall r \geq \text{step}_{i}(P), \exists P \in Val \) such that \( P[O'/x_1,\ldots, O'/x_n] \rightarrow_{r}^{*} P \), and \( Q[O'/x_1,\ldots, O'/x_n] \rightarrow_{r}^{*} Q \) (by Lem. 4.7). Furthermore, \( P \Downarrow_{a}^{0} \) if and only if \( P \Downarrow_{i} \) and \( \text{step}_{i}[P] = \text{step}_{a}^{0}[P] \).

The proof is carried out by induction on the derivation of \( M \Downarrow_{a}^{0} N \).

\[ \ast \theta: \text{The proof follows immediately from the induction hypothesis.} \]

**var:** Let

\[ m \geq 0 \quad M_{i} \Downarrow_{1}^{0} N_{i} \quad xM_{1}\ldots M_{m} \Downarrow_{a}^{0} xN_{1}\ldots N_{m} \quad (\text{var}). \]
By Lemma 4.8 \( \forall r_j \geq \text{step}_h(M_j) \) \( M_j[O^{r_1}/x_1, ..., O^{r_j}/x_n] \rightarrow^* M_j \in \text{Val} \) (1 \( \leq j \leq m \)). Let \( \pi = \text{step}_h^0(M) = 1 + \text{step}_h^0(M_1) + ... + \text{step}_h^0(M_m) \). Let \( \rho > \max\{n, \pi\} \). Since \( \pi > m \), then

\[
(\lambda x_1...x_n.x_1M_1...M_m)O^\rho...O^\rho \rightarrow^*_\rho O^\rho M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho O^h
\]

for some \( h \geq 0 \), where \( M'_1 \equiv M_1[O^\rho/x_1, ..., O^\rho/x_n] \).

**head:** Let the last used rule be

\[
\frac{Q \Downarrow^1 R \quad P[R/z]M_1...M_m \Downarrow^0 N}{(\lambda z.P)RM_1...M_m \Downarrow^1 N} \quad \text{(head)}.
\]

If \( \pi = \text{step}_h^0(P[R/z]M_1...M_m) \), then by induction, \( \forall r > \max\{n, \pi\}, \exists h' \geq 0: \)

\[
(\lambda x_1...x_n.P[R/z]M_1...M_m)O^\rho...O^\rho \rightarrow^*_\rho O^{h'}.
\]

Let \( \rho > \max\{n, 1 + \text{step}_h^1(Q) + \pi\} \). So by Lemma 4.8 \( \exists Q' \in \text{Val}^0 \) such that both

\[
Q' \equiv Q[O^\rho/x_1, ..., O^\rho/x_n] \rightarrow^*_\rho Q'
\]

and

\[
R' \equiv R[O^\rho/x_1, ..., O^\rho/x_n] \rightarrow^*_\rho Q'.
\]

Let \( P' \equiv P[O^\rho/x_1, ..., O^\rho/x_n] \) and \( M'_1 \equiv M_1[O^\rho/x_1, ..., O^\rho/x_n] \). Then

\[
(\lambda x_1...x_n.P[R/z]M_1...M_m)O^\rho...O^\rho \equiv P'[R'/z]M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho
\]

\[

\rightarrow^*_\rho P'[Q'/z]M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho O^{h''} \quad (h'' \geq 0)
\]

by Church-Rosser, and finally

\[
(\lambda x_1...x_n.(\lambda z.P)QM_1...M_m)O^\rho...O^\rho \rightarrow^*_\rho (\lambda z.P')Q'M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho
\]

\[

\rightarrow^*_\rho (\lambda z.P')Q'M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho P'[Q'/z]M'_1...M'_m O^\rho...O^\rho \rightarrow^*_\rho O^{h''}.
\]

\( \square \)

Now we are able to prove our result.

**Theorem 5.5** (\( v \)-solvability). \( M \Downarrow^0 \) if and only if \( M \) is \( v \)-solvable.

**Proof.** Let \( FV(M) = \{x_1, ..., x_n\} \).
\[
(\lambda x_1 \ldots x_n . M)^{O^r \ldots O^r} \searrow^* O^h \quad (h \geq 0).
\]

So \((\lambda x_1 \ldots x_n . M)^{O^r \ldots O^r} R_1 \ldots R_h \searrow^* I\), for all \(R_1, \ldots, R_h \in \text{Val}\).

\[\Rightarrow\]: By Lemma 5.4 there are \(r\) and \(h\) such that:

By Lemma 5.3.

It can be interesting to compare the notions of \(v\)-normal form, valuable term and \(v\)-solvable term. \(\lambda x.(\lambda y.\Delta)(xI)\Delta\) and \(\lambda x.\Delta\Delta\) are respectively a \(\beta_v\)-normal-form and a value, and are both \(v\)-unsolvable.

We can classify the \(v\)-unsolvable terms as follows.

**Definition 5.6.** Let \(P\) be \(v\)-unsolvable. \(P\) is of order 0 if and only if there is no \(Q\) such that \(P \equiv \lambda x.Q\). \(P\) is of order \(k + 1\) if \(P \equiv \lambda x.Q\) and \(k\) is the maximum integer such that \(Q\) is \(v\)-unsolvable of order \(k\), while it is of infinite order if this integer does not exists.

All the \(v\)-unsolvable terms of order 0 can be consistently equated (see [4]). Moreover the relation between potentially valuable and \(v\)-solvable terms can be now stated as follows.

**Proposition 5.7.** A term is not potentially valuable if and only if it is \(v\)-unsolvable of order 0.

A \(\lambda\beta_v\)-theory is a congruence relation on terms closed under the \(\beta_v\)-equality. Let us recall that the \(\lambda\)-theories can be classified into sensible and semi-sensible, the former being these equating all unsolvable terms, and the latter these never equating a solvable term to an unsolvable one. We will introduce a similar classification for the \(\lambda\beta_v\)-theories.

**Definition 5.8.** i) A \(\lambda\beta_v\)-theory is \(v\)-sensible if and only if it equates all the \(v\)-unsolvable terms.

ii) A \(\lambda\beta_v\)-theory is \(v\)-semi-sensible if and only if it never equates a \(v\)-solvable term to a \(v\)-unsolvable one.

The \(\lambda\beta_v\)-theory induced by the seed operational equivalence is \(T_{\text{seed}} = \{(M, N) | M \approx_v N\}\). It immediate to see that:

**Proposition 5.9.** \(T_{\text{seed}}\) is not \(v\)-sensible.

**Proof.** Consider the two terms \(\Delta\Delta\) and \(\lambda x.\Delta\Delta\). They are both \(v\)-unsolvable but the former is not valuable while the latter is a value. \(\Box\)

Now we will prove that \(T_{\text{seed}}\) is not \(v\)-semi-sensible.

Let \(Y_v \equiv (\lambda xf.f.(\lambda z.xxfz))(\lambda xf.f.(\lambda z.xxfz)), A \equiv \lambda yz.(\lambda uv.xuv)(yz)\) and \(R \equiv Y_v AI\).

It is immediate to check that, for all \(M \in \text{Val}, Y_v M \searrow^* M(\lambda z.Y_v M z)\). The combinator \(Y_v\) is a recursion combinator in the call-by-value setting. We will prove
that $R$ and $I$ have the same operational behaviour. The next lemma will allow us to consider just contexts of a particular shape.

**Lemma 5.10.** Let $M, N \in \Lambda$ and let $FV(M) \cup FV(N) \subseteq \{x_1, \ldots, x_n\}$.

$s$ is not solvable if and only if there exist $C[\ ] \equiv [\ ]M_1 \ldots M_m$ such that $C[M], C[N] \in \Lambda^0$ and $C[M] \downarrow_s$ or vice versa, for some $M_1, \ldots, M_m \in \Lambda$.

**Proof.** See [4] (Th. 33).

**Lemma 5.11.** Let $P_1, \ldots, P_k \in \Lambda^0$ $(k \geq 1)$.

$IP_1 \ldots P_k \downarrow_s M$ if and only if $RP_1 \ldots P_k \downarrow_s \lambda v. (\lambda z. Y_v A z) M v$.

**Proof.** By induction on $k$.

$k = 1$: $IP_1 \downarrow_s M$ if and only if $P_1 \downarrow_s M$ if and only if $RP_1 \downarrow_s \lambda v. (\lambda z. Y_v A z) M v$ (by the secd-rules and $R \downarrow_s \lambda x. (\lambda u v. (\lambda z. Y_v A z) u v)(1x)$).

$k > 1$: By induction $IP_1 \ldots P_{k-1} \downarrow_s N$ if and only if $RP_1 \ldots P_{k-1} \downarrow_s \lambda v. (\lambda z. Y_v A z) N v$.

$IP_1 \ldots P_k \downarrow_s M$ if and only if $N \equiv \lambda y. N'$ and $P_k \downarrow_s M'$ and $N'[M'/y] \downarrow_s M$ if and only if $((\lambda z. Y_v A z) N v)[M_k/v] \downarrow_s \lambda v. (\lambda z. Y_v A z) M v$ (by secd-rules) if and only if $RP_1 \ldots P_k \downarrow_s \lambda v. (\lambda z. Y_v A z) M v$.

**Theorem 5.12.** $T_{\text{seed}}$ is not $v$-semi-sensible.

**Proof.** We will prove that $R \simeq I$. Since $R, I \in \Lambda^0$, by Lemma 5.10 we can consider just contexts of the shape $C[\ ] \equiv [\ ]M_1 \ldots M_m$. If $m = 0$ then the secd machine stops for both $I$ and $R$. Otherwise the proof follows from Lemma 5.11.

## 6. Logical Characterization

In this section we will present a type assignment system which allows a complete characterization of the $v$-solvable terms.

**Definition 6.1.** Let $\nu$ and $\alpha$ be two type constants. Let $T$ be the set of types $\sigma$ built out from the following grammar:

$\sigma ::= \nu | \alpha | \sigma_1 \cap \ldots \cap \sigma_n \rightarrow \sigma \quad (n \geq 1)$.

$T$ will be ranged over by $\sigma, \tau, \pi, \rho, \mu, \ldots$.

The $\rightarrow$ type-constructor is associative on the right and the intersection type-constructor $\cap$ binds stronger than $\rightarrow$. The types are considered modulo permutations of types bound by intersection constructor.

All types have the following shape:

$\sigma_1^1 \cap \ldots \cap \sigma_n^1 \rightarrow \sigma_1^2 \cap \ldots \cap \sigma_n^2 \rightarrow \ldots \rightarrow \sigma_1^m \cap \ldots \cap \sigma_n^m \rightarrow \rho$
for some $m, n$ where $\rho$ is either $\nu$ or $\alpha$. In the latter case the type is named proper. Let a proper type be denoted by $\sigma_p$ and the subset of proper type by $T_p$. In the rest of the paper, we will use $\equiv$ for denoting the syntactical identity both on terms and types.

**Definition 6.2.**

i) Let a basis be a finite set of assignments of the shape $x : \sigma$, where $x$ is a variable and $\sigma$ is a type. If $B$ is a basis, let $\text{dom}(B) = \{x \mid x : \sigma \in B\}$.

ii) The following type assignment system proves statements of the shape: $B \vdash M : \sigma$ where $B$ is a basis, $M \in \Lambda$ and $\alpha \in T$. The rules are:

$$\frac{}{\{x : \sigma\} \vdash x : \sigma} \quad (\text{var})$$

$$\frac{}{\emptyset \vdash \lambda x. M : \nu} \quad (\nu)$$

$$\frac{B \vdash M : \tau \quad x \notin \text{dom}(B)}{B \vdash \lambda x. M : \nu \rightarrow \tau} \quad (\rightarrow_\nu \ I)$$

$$\frac{B \vdash M : \tau \quad x \notin \text{dom}(B)}{B \vdash \lambda x. M : \nu \rightarrow \tau} \quad (\rightarrow \ I)$$

$$\frac{B \vdash M : \tau \quad B \cup \{x : \sigma_1, \ldots, x : \sigma_n\} \vdash M : \tau \quad x \notin \text{dom}(B)}{B \vdash \lambda x. M : \sigma_1 \cap \ldots \cap \sigma_n \rightarrow \tau} \quad (\rightarrow \ I)$$

$$\frac{B \vdash M : \sigma_1 \cap \ldots \cap \sigma_n \rightarrow \tau \quad (B_i \vdash N : \sigma_i)_{1 \leq i \leq n} \quad (\rightarrow \ E)}{B \cup_{i=1}^{\nu} B_i \vdash MN : \tau}$$

We will denote by $D : B \vdash M : \sigma$ a derivation $D$ proving $B \vdash M : \sigma$; and by $D^\prime \subseteq D$ the fact that $D^\prime$ is a subderivation of $D$.

**Proposition 6.3.**

i) Let $M \in C$. Then $\exists B. \sigma. \ B \vdash M : \sigma$. In particular, $\exists B. \ B \vdash M : \nu$.

ii) Let $M \in VH$. Then $\exists B. \sigma_p. \ B \vdash M : \sigma_p$.

iii) Let $M \in Val$. Then $\exists B. \ B \vdash M : \nu$.

**Proof.**

i) By induction on $M$. If $M$ is an abstraction apply directly the rule $(\nu)$. If $M \equiv xM_1 \ldots M_m$, by induction $\exists B_i. \sigma_i$. $B_i \vdash M_i : \sigma_i$ ($1 \leq i \leq m$); then for every $\sigma$, there is a derivation using one application of rule $(\nu)$ followed by $m$ applications of rule $(\rightarrow \ E)$, proving:

$$\{x : \sigma_1 \rightarrow \ldots \rightarrow \sigma_m \rightarrow \sigma\} \cup_i B_i \vdash M : \sigma.$$  

Note that in particular it can be $\sigma \equiv \nu$.

ii) By induction on $M$. If $M \equiv xM_1 \ldots M_m$, take the proof of point i) and choose $\sigma$ to be proper. The case $M$ is an abstraction is trivial by induction.

iii) Trivial by $(\nu)$ rules.

$\square$

We will prove that the typability in the above type assignment system is preserved by $\beta$-reduction and by a particular case of $I$-expansion. In particular, since $\rightarrow_i$ implies $\rightarrow_\beta$, and the $i$-expansion implies the expansion we considered, it turns out that the system is closed under $=_i$. 

Lemma 6.4 (Subject reduction). $B \vdash M : \sigma$ and $M \rightarrow^* \beta M'$ imply $\exists B'$ such that $B' \vdash M' : \sigma$.

Proof. The proof is by induction on the number $m$ of $\beta$-reduction steps. The case $m = 0$ is obvious. Otherwise, let $D : B \vdash M : \sigma$. We will build a new derivation $D'$ proving $B' \vdash M' : \sigma$, where $M \rightarrow^* \beta M' \rightarrow^*_\beta M'^*$. Then the result follows by induction.

Let $M \equiv C[(\lambda x.P)Q]$ and $M' \equiv C[P[Q/x]]$. If $(\lambda x.P)Q$ occurs in a subterm of $M$ typed using the rule $(\nu)$, then $D'$ can be obtained by replacing $(\lambda x.P)Q$ by $P[Q/x]$ in all subjects of $D$. Otherwise there are two cases.

1) There is a subderivation $S \subseteq D$ of the shape:

\[
\begin{array}{c}
\vdots \\
(\rightarrow I) \quad B \uplus \{x : \sigma_1, \ldots, x : \sigma_n\} \vdash P : \tau \\
(\rightarrow E) \quad B \vdash \lambda x.P : \sigma_1 \cap \ldots \cap \sigma_n \rightarrow \tau \\
\vdots \\
B \vdash Q : \sigma_{i(1 \leq i \leq n)} \\
\hline \\
\end{array}
\]

Then by:

a) replacing the $i$-th application of the $(\text{var})$-rule typing $x$ by $D_i$ $(1 \leq i \leq n)$;
b) replacing every occurrence of $x$ in the subjects by $Q$;
c) replacing every assignment $x : \sigma_i$ by the assignments in $B_i$;
d) erasing the rules $(\rightarrow I)$ and $(\rightarrow E)$;

the following subderivation $S'$ can be built.

\[
\begin{array}{c}
\vdots \\
D_1 \\
\vdots \\
\vdots \\
\vdots \\
\hline \hline
B_1 \vdash Q : \sigma_1 \\
B_n \vdash Q : \sigma_n \\
\hline \\
B' \cup_{i=1}^n B_i \vdash P[Q/x] : \tau
\end{array}
\]

Note that $B' = B$. The desired $D'$ is then obtained by replacing in $D$ the subderivation $S$ by $S'$ and finally, by replacing in the rest of derivation $(\lambda x.P)Q$ by $P[Q/x]$.

2) Let the redex be introduced by an application of the rule $(\rightarrow \nu I)$ followed by an application of the rule $(\rightarrow E)$. In this case the proof is similar, but it is possible...
that \( B' \neq B \) since either \( x \) does not occur in \( P \) or it occurs just in subterms of \( P \) typed by \( \nu \).

**Lemma 6.5.** Let \( D : B \vdash P[Q/x] : \sigma \). If \( Q \in \mathcal{C} \) then \( \exists B' \) such that \( B' \vdash (\lambda x.P)Q : \sigma \).

**Proof.** The occurrences of \( Q \) considered for the expansion in \( P \) can be divided in two groups: let \( Q_1, \ldots, Q_q \) \((q \geq 0)\) be those occurrences of \( Q \) such that there is \( \mathcal{D}_i : B_i \vdash Q_i : \sigma_i \) and \( \mathcal{D}_i \subseteq D \) \((1 \leq i \leq q)\) and let \( Q_{q+1}, \ldots, Q_{q+p} \) \((p \geq 0)\) be those occurrences of \( Q \) which are not typed by subderivations of \( D \) \( \langle i,e \rangle \) these occurrences are in subterms of \( P \) typed by the constant \( \nu \). We will consider two cases, according to \( q = 0 \) or \( q > 0 \).

\( q = 0 \) Every occurrence of \( Q \) in \( P[Q/x] \) occurs in a subterm of \( P \) typed by a rule \( \langle \nu \rangle \).

Since, by hypothesis, \( Q \in \mathcal{C} \), by Property 6.3, there is \( B^* \) and a derivation \( S^* \) proving \( B^* \vdash Q : \nu \).

Let replace in \( D \) every such occurrence of \( Q \) by \( x \) \( \langle note \ that \ x \ is \ not \ typed \rangle \); the result is a derivation \( D' : B \vdash P : \sigma \), where \( x \notin \text{dom}(B) \).

Thus, \( D \) is the following subderivation:

\[
\begin{array}{ccc}
D' & : & S^* \\
\vdash B \vdash P : \sigma & x \notin \text{dom}(B) & \\
\vdash B \vdash \lambda x.P : \nu \rightarrow \sigma & & \vdash B^* \vdash Q : \nu \\
\end{array}
\]

\( B \cup B^* \vdash (\lambda x.P)Q : \sigma \)

\( q > 0 \) Let \( S \) be the subderivation obtained from \( D \) by:

a) replacing every \( \mathcal{D}_i : B_i \vdash Q : \sigma_i \) by the rules

\[
\begin{array}{c}
\{x : \sigma_i\} \vdash x : \sigma_i^{(\text{var})} \quad (1 \leq i \leq q)
\end{array}
\]

b) replacing every rule \( \vdash B_i[Q/x] : (\nu) \) by \( \vdash B_i[Q/x] : (\nu) \), where \( P_i \) are the subterms of \( P \) containing the \( i \)-th untyped occurrence of \( Q \) \((q < i \leq q + p)\)

c) replacing every occurrence of \( Q \) in the subjects by \( x \) and adjusting elsewhere the basis.

The result is \( S : B' \cup \{x : \sigma_1, \ldots, x : \sigma_q\} \vdash P : \tau \) where

\[
B' = \left\{ z : \mu \mid z : \mu \in B, z \in \text{FV}(M) \right\}
\]

the rule \( \{z : \mu\} \vdash z : \mu^{(\text{var})} \) occurs out of the \( \mathcal{D}_i \) \((1 \leq i \leq q)\).
The subderivation $D$ is obtained by adjoining an application of ($\rightarrow E$) and ($\rightarrow I$) rules to $S$, in the following way:

\[
\begin{array}{c}
S \\
\vdots \\
\vdots \\
\vdots \\
\vdots \\
\vdots \\
\hline \\
B' \cup \{x : \sigma_1, \ldots, x : \sigma_n\} \vdash P : \tau \\
B' \vdash \lambda x.P : \sigma_1 \cap \ldots \cap \sigma_q \rightarrow \tau \\
B_i \vdash Q : \sigma_i(1 \leq i \leq n) \\
\hline \\
B' \cup_{i=0}^{\#} B_i \vdash (\lambda x.P)Q : \tau
\end{array}
\]

Note that $B' \cup_{i=0}^{\#} B_i = B$.

Lemma 6.6 (Subject expansion). $B \vdash M : \sigma$ and $M \rightarrow_i^* M$ imply $\exists B$ such that $B \vdash M : \sigma$.

Proof. By induction on the number $m$ of $i$-reduction steps. The case $m = 0$ is obvious. Let consider the case $m = 1$; then the general case follows directly from the induction hypothesis.

Let $M \equiv C[P(Q/x)]$ and $\bar{M} \equiv \bar{C}[(\lambda x.P)Q]$. Since $\bar{M} \rightarrow_i M$, then $P(Q/x)$ does not occur under an abstraction. This means that, if $D : B \vdash M : \sigma$, then $P(Q/x)$ cannot occur in a subterm of $M$ typed by the rule ($\nu$), so there is a subderivation $S \subseteq D$, such that $S : B^{*} \vdash P(Q/x) : \tau$, for some $B^{*} \vdash \lambda x.P)Q : \tau$. Then, by Lemma 6.5, there is $S' : B' \vdash P[(\lambda x.P)Q] : \tau$. The conclusion is obtained from $D$ by replacing $S$ by $S'$, by adjusting the basis and by replacing every occurrence of $P(Q/x)$ in the subjects in $D$ by $(\lambda x.P)Q$.

Theorem 6.7 (Characterization of potentially solvable terms). $M \not\rightarrow_i N$ if and only if $\exists B$. $B \vdash M : \sigma$.

Proof. ($\Leftarrow$): $M \not\rightarrow_i N$ if and only if $M \rightarrow_i^* N$ and $N \in C$. Then the proof follows by induction on the length of the reduction from $M$ to $N$, using Proposition 6.3.1 and Lemma 6.4 (since $M \rightarrow_i N$ implies $M \rightarrow_i N$).
($\Rightarrow$): See the Appendix B.

Theorem 6.8 (Characterization of $\nu$-solvable terms). $M \not\rightarrow_\nu N$ if and only if $\exists B$. $B \vdash M : \sigma_p \in T_p$.

Proof. ($\Leftarrow$): By Proposition 5.2, $M \not\rightarrow_\nu^0 N$ implies $M \rightarrow_i^* N$, for some $N$. Then the proof can be carried out by induction on the number of steps of the reduction from $M$ to $N$, using Proposition 6.3.1 and Lemma 6.4 (since $M \rightarrow_i^* N$ implies $M \rightarrow_i^* N$).
Lemma 7.1. $M \vdash^v N$ and $(N)$ is defined imply $(M)$ is defined.

Proof. $M \vdash^v N$ means that $M \equiv C[(\lambda x.P)Q]$ and $N \equiv C[P/Q/x]$. The proof is given by induction on $(N)$. $(C[P/Q/x])$ is defined implies that $C[\ ]$ must have one of the following shapes:

1. $C[\ ] \equiv [ ]P_1\ldots P_n$;
2. $C[\ ] \equiv (\lambda y.C'[\ ])(P_1\ldots P_n)$;
3. $C[\ ] \equiv (\lambda y.P_0)P_1\ldots P_{n-1}C'[\ ]P_n$ ($n > 0, 1 \leq i \leq n$).

Let recall that $(\lambda x.P)Q \vdash^v P/Q/x$ implies $Q$ is a value and so $(Q) = 0$. We can assume, without loss of generality, that $x \not\in \text{FV}(P_i)$, for all $i$ ($1 \leq i \leq n$); otherwise we can rename the bound variable $x$ with a fresh variable.

Let $C[\ ]$ be of the shape 1. Then $(C[P/Q/x])$ defined implies that $P/Q/x$ is of the shape $M_0M_1\ldots M_n$ ($m \geq 0$), with $M_0 \equiv (\lambda z.M_0')$ or $M_0 \equiv z$ and $m = 0$.

Let $m = 0$ and $n = 0$. Then $C[P/Q/x] \equiv M_0 \in \text{Val}$. Then $(C[P/Q/x]) = 0$, and $(C[(\lambda x.P)Q]) = 1 + (Q) + (P_1) = 1 + 0 + 0 = 1$.

Let $m > 0$, and $n \geq 1$. Then $C[P/Q/x] \equiv M_0M_1\ldots M_n$ and $M_0 \equiv (\lambda z.M_0')$, so by hypothesis $(M_0M_1\ldots M_nP_1\ldots P_n)$ is defined. Furthermore, since $(Q) = 0$, $(N) = ((\lambda x.(\lambda z.M_0)M_1\ldots M_n)QP_1\ldots P_n) = 1 + (Q) + (\lambda z.M_0)M_1\ldots M_nP_1\ldots P_n$ is defined.

The case $m = 0$ and $n \geq 1$ is similar to the previous one, but simpler.

Let $C[\ ]$ be of the shape 2. $(C[P/Q/x]) = 1 + (P_1) + (C'[(\lambda x.P)Q][P_1/y]P_2\ldots P_n)$. By induction $(C'[(\lambda x.P)Q][P_1/y]P_2\ldots P_n)$ is defined, and we are done, since by definition $(C[(\lambda x.P)Q]) = 1 + (P_1) + (C'[(\lambda x.P)Q]$ 

$[P_1/y]P_2\ldots P_n)$.

Let $C[\ ]$ be of the shape 3.

Let $i = 1$. So $(C[P/Q/x]) = 1 + (C'P/[x/x]) + (P_0C'[P/Q/x]/y)P_2\ldots P_n)$. By induction $(C'[(\lambda x.P)Q])$ and $(P_0C'[P/Q/x]/y)P_2\ldots P_n)$ are both defined and we have done, since $(C[(\lambda x.P)Q]) = 1 + (C'[(\lambda x.P)Q]$ 

$[P_0C'[P/Q/x]/y]P_2\ldots P_n)$.
Let $i > 1$. Then

\[ \langle C[PQ/x]\rangle = 1 + \langle P_1 \rangle + \langle P_0[P_1/y]P_2 \ldots P_{i-1}[C'[PQ/x]/y]P_{i+1} \ldots P_n \rangle, \]

and by induction $\langle P_0[P_1/y]P_2 \ldots P_{i-1}[C'[(\lambda x.P)Q]/y]P_{i+1} \ldots P_n \rangle$ is defined. So we have done, since by definition

\[ \langle C[(\lambda x.P)Q]\rangle = 1 + \langle P_1 \rangle + \langle P_0[P_1/y]P_2 \ldots P_{i-1}[C'[(\lambda x.P)Q]/y]P_{i+1} \ldots P_n \rangle. \]

Lemma 7.2. $\langle M \rangle$ is defined and $M \rightarrow^* N$ imply $\langle M \rangle \geq \langle N \rangle$.

Proof. Let $\langle M \rangle = k$. The proof is given by induction on the following pair: $(k, p)$, where $p$ is the numbers of steps of the reduction $M \rightarrow^* N$, ordered according to the lexicographical order. The cases where either $\langle M \rangle = 0$ or $p = 0$ are trivial.

Let the reduction path be: $M \rightarrow_\beta R_1 \rightarrow_\beta \ldots \rightarrow_\beta R_p \equiv N$ ($p \geq 0$). Clearly $M \equiv (\lambda x.M_0)M_1 \ldots M_m$, so let $h' = \langle M_1 \rangle$, $h'' = \langle M_0[M_1/x]M_2 \ldots M_m \rangle$ and so $k = 1 + h' + h''$. Let $p = 1$. There are three cases:

1. If $R_1 \equiv M_0[M_1/x]M_2 \ldots M_m$ then $\langle R_1 \rangle = h'' < k$. Thus the proof is trivial.

2. Let $R_1 \equiv (\lambda x.N_0)M_1N_2 \ldots N_m$ where $\exists j$ (unique) $M_j \rightarrow_\beta N_j$, while for $i \neq j$ $M_i \equiv N_i$ (0 $\leq i,j \leq m$ and $i,j \neq 1$). Clearly $M_0[M_1/x]M_2 \ldots M_m \rightarrow_\beta N_0[M_1/x]N_2 \ldots N_m$; thus $h'' < k$ implies, by induction $\langle N_0[M_1/x]N_2 \ldots N_m \rangle \leq h''$. Finally $\langle R_1 \rangle = 1 + \langle M_1 \rangle + \langle N_0[M_1/x]N_2 \ldots N_m \rangle \leq k$ and the proof is done.

3. Let $R_1 \equiv (\lambda x.M_0)N_1M_2 \ldots M_m$, where $M_1 \rightarrow_\beta N_1$. By induction $\langle M_1 \rangle \geq \langle N_1 \rangle$. Furthermore $h'' < k$ and $M_0[M_1/x]M_2 \ldots M_m \rightarrow^*_\beta M_0[N_1/x]M_2 \ldots M_m$ imply, by induction, $\langle M_0[M_1/x]M_2 \ldots M_m \rangle \leq h''$. Thus the conclusion follows, trivially, by definition of weight.

Since $(\langle M \rangle, p)$ is greater than $(\langle R_1 \rangle, p-1)$, the complete proof follows by induction.

Now we are able to prove the Proposition 4.5.

Proof. i) By induction on the number of steps of the reduction $M \rightarrow^*_\nu \tilde{M} \in Val$ using Lemma 7.1.

ii) By i) and by Lemma 7.2.

iii) By ii), since the $\beta_\nu$-reduction is a special case of the $\beta$-reduction.

8. Appendix B

Proof. ($\Rightarrow$) of Theorem 6.7.

The proof will be given by a computability argument.

Let define the following predicate: $P(\sigma, M) \Rightarrow$ there is a basis $B$ such that $B \vdash M : \sigma$ and $M \not\approx_1$.

Let $\tilde{M}$ denote a sequence of terms $M_1 \ldots M_m$, for some $m \geq 0$. 
Proposition 8.1.

i) \( P(\sigma_1^i \ldots \sigma_n^i \rightarrow \tau, x \vec{M}) \) and \( P(\sigma_i, N) \) imply \( P(\tau, x \vec{M}N) \) (\( 1 \leq i \leq n \)).

ii) \( P(\tau, Mx) \) and \( x \notin \text{FV}(M) \) imply \( P(\sigma_1 \ldots \sigma_n \rightarrow \tau, M) \), for some \( \sigma_1, \ldots, \sigma_n \).

Proof.

i) Clearly \( \exists B, B' \) such that \( B \vdash x \vec{M} : \sigma_1 \ldots \sigma_n \rightarrow \tau \) and \( B' \vdash N : \sigma_i \) imply \( B \cup B' \vdash x \vec{M}N : \tau \). Let us prove that if \( x \vec{M} \downarrow_i \) and \( N \downarrow_i \) then \( x \vec{M}N \downarrow_i \).

\( x \vec{M} \downarrow_i \) implies there is a derivation whose last applied rule is:

\[
\frac{M_j \downarrow_i M'_j \ (1 \leq i \leq m)}{xM_1 \ldots M_m \downarrow_i xM'_1 \ldots M'_m}.
\]

Since \( N \downarrow_i \) the following derivation can be built:

\[
\frac{M_j \downarrow_i M'_j \ (1 \leq i \leq m) \ N \downarrow_i N'}{xM_1 \ldots M_mN \downarrow_i xM'_1 \ldots M'_mN'}.
\]

\( P(\tau, x \vec{M}N) \) follows by rule \((\rightarrow E)\).

ii) The proof is given by induction on the derivation proving \( Mx \downarrow_i \). The only not trivial case is when the last applied rule is:

\[
\frac{Q \downarrow_i R \ P[R/z]M_1 \ldots M_m \downarrow_i N}{(\lambda z. P)QM_1 \ldots M_m x \downarrow_i N}
\]

where \( M \equiv (\lambda z. P)QM_1 \ldots M_m \). By induction \( P[R/z]M_1 \ldots M_m \downarrow_i N \) implies \( P[R/z]M_1 \ldots M_m \downarrow_i \), so \( M \downarrow_i \). Moreover, for some basis \( B, B_i \) (\( 1 \leq i \leq n \)), there is a derivation ending with a rule:

\[
\frac{B \vdash M : \sigma_1 \ldots \sigma_n \rightarrow \tau \ (B_i \vdash x : \sigma_i)_{1 \leq i \leq n}}{B \cup_i B_i = B \cup \{x : \sigma_1, \ldots, x : \sigma_n\} \vdash Mx : \tau \ (\rightarrow E)}
\]

Since \( x \notin \text{FV}(M) \), it must be \( x \notin \text{dom}(B) \). Then \( P(\sigma_1 \ldots \sigma_n \rightarrow \tau, M) \).

Note that the case \( \sigma \equiv \nu \rightarrow \tau \), is implicitly considered.

Now let define the following computability predicate:

- \( \text{Comp}^i(\alpha, M) \Leftrightarrow P(\alpha, M) \).
- \( \text{Comp}^i(\nu, M) \Leftrightarrow P(\nu, M) \).
- \( \text{Comp}^i(\sigma_1 \ldots \sigma_n \rightarrow \tau, M) \Leftrightarrow \forall j \ (1 \leq j \leq n) \ (\text{Comp}^i(\sigma_j, N) \implies P(\pi(\tau, MN)) (1 \leq j \leq n) \).

Lemma 8.2. \( \text{Comp}^i(\sigma, M[N/x]) \) and \( N \in \mathcal{C} \) imply \( \text{Comp}^i(\sigma, (\lambda x. M)N) \).

Proof. By induction on \( \sigma \). The basis case follows from the definition of \( P \) and by Lemma 6.5. The general case follows immediately from the induction hypothesis.

We will prove: \( B \vdash M : \sigma \Rightarrow \text{Comp}^i(\sigma, M) \Rightarrow P(\sigma, M) \Rightarrow M \downarrow_i \).
Lemma 8.3. \( \text{i) } \mathcal{P}(\sigma, x\bar{M}) \Rightarrow \text{Comp}^i(\sigma, x\bar{M}). \)
\( \text{ii) } \text{Comp}^i(\sigma, M) \Rightarrow \mathcal{P}(\sigma, M). \)

Proof. By mutual induction on \( \sigma \). The only not trivial case is \( \sigma \equiv \mu_1 \ldots \mu_m \rightarrow \tau. \)
\( \text{i) } \text{Comp}^i(\mu_j, N) \Rightarrow \mathcal{P}(\mu_j, N) \text{ by induction on ii) } (1 \leq j \leq m). \)
\( \mathcal{P}(\mu_1 \ldots \mu_m \rightarrow \tau, x\bar{M}) \text{ and } \mathcal{P}(\mu_1, N) \Rightarrow \mathcal{P}(\tau, x\bar{M}N) \text{ by Property 8.1.1. } \)
\( \Rightarrow \text{Comp}^i(\tau, x\bar{M}) \text{ (**) by induction. Finally (**) and (***) together imply } \text{Comp}^i(\mu_1 \ldots \mu_m \rightarrow \tau, x\bar{M}), \text{ by definition of Comp}^i. \)
\( \text{ii) } \text{Let } x \not\in \text{FV}(M). \) Clearly \( \mathcal{P}(\mu_j, x) \Rightarrow \text{Comp}^i(\mu_j, x) \) (1 \( \leq j \leq m) \) by induction on \( \text{i) } \text{Comp}^i(\mu_1 \ldots \mu_m \rightarrow \tau, M) \) and \( \text{Comp}^i(\mu_j, x) \) (1 \( \leq j \leq m) \) \( \Rightarrow \text{Comp}^i(\tau, Mx) \), by definition of \( \text{Comp}^i \Rightarrow \mathcal{P}(\tau, Mx) \) by induction \( \Rightarrow \mathcal{P}(\mu_1 \ldots \mu_m \rightarrow \tau, M) \) by Property 8.1.2. \]

\[ \square \]

Lemma 8.4. Let \( \text{FV}(M) = \{x_1, \ldots, x_n, x_{n+1}, \ldots, x_{n+k}|n, k \geq 0\}, \) \( B = \{x_j : \sigma_j^n|1 \leq j \leq n, 1 \leq r \leq m_j, 1 \leq m_j\} \) and \( D : B \vdash M : \tau. \)
\( \text{If } Q_j \in \mathcal{C} \text{ (1} \leq j \leq n + k \text{) and } \text{Comp}^i(\sigma_j, Q_h) \text{ (1} \leq r \leq m_j, 1 \leq h \leq n) \text{ then } \text{Comp}^i(M(Q_1/x_1, \ldots, Q_{n+k}/x_{n+k})). \)

Proof. By induction on the derivation. If the last applied rule is either \( (\varnothing) \text{ or } (\nu) \) the proof is trivial. If the last applied rule is \( (\rightarrow E) \) the proof follows by induction, the definition of computability and \( (\rightarrow E) \). Let consider \( (\rightarrow I). \) It must be \( M \equiv \lambda y.P \) and \( \tau \equiv \mu_1 \ldots \mu_p \rightarrow \tau'. \) so
\[ \begin{array}{c}
B \cup \{y : \mu_1, \ldots, y : \mu_p\} \vdash P : \tau'\quad y \not\in \text{dom}(B)
\end{array} \]
\( \frac{B \vdash (\lambda y.P) : \mu_1 \ldots \mu_p \rightarrow \tau'}{(\rightarrow I).} \)

Let assume \( N \in \mathcal{C} \) and \( \text{Comp}^i(\mu_j, N) \) (1 \( \leq j \leq p). \) This implies (by Lem. 8.3.2) that there are basis \( B_j \) such that \( B_j \vdash N : \mu_j \) (1 \( \leq j \leq p). \) By induction
\( \text{Comp}^i(\tau', P[Q_1/x_1, \ldots, Q_{n+k}/x_{n+k}, N/y]). \)

By Lemma 8.2:
\[ \text{Comp}^i(\tau', (\lambda y.P[Q_1/x_1, \ldots, Q_{n+k}/x_{n+k}])N) \]
which together with \( \text{Comp}^i(\mu_j, N), \) implies
\[ \text{Comp}^i(\mu_1 \ldots \mu_p \rightarrow \tau', M[Q_1/x_1, \ldots, Q_{n+k}/x_{n+k}]). \]
The case \( (\rightarrow \nu, I) \) is similar, using Proposition 6.3.1. \[ \square \]

Now we are able to conclude the proof.

Let \( B \vdash M : \sigma, \text{FV}(M) = \{x_1, \ldots, x_n, x_{n+1}, \ldots, x_{n+k}\} \) and \( B = \{x_j : \sigma_j^n|1 \leq r \leq m_j, 1 \leq j \leq n, 1 \leq m_j\}. \)

By Lemma 8.3i), \( \text{Comp}^i(\sigma_j, x_j) \) (1 \( \leq r \leq m_j, 1 \leq j \leq n). \) Then by Lemma 8.4 \( \text{Comp}^i(\sigma, M), \) which implies \( \mathcal{P}(\sigma, M) \) (by Lem. 8.3.2). \[ \square \]
Proof. (⇒) part of the Theorem 6.8.

The proof is very similar to the previous one. Let \( \sigma_p \in T_p \) and define the following predicate:

\[ R(\sigma_p, M) \iff B \vdash M : \sigma_p, \text{ for some basis } B, \text{ and } M \not\vdash_\eta. \]

Proposition 8.5.

1. \( R(\sigma_1 \cap \ldots \cap \sigma_n \Rightarrow \tau_p, x \bar{M}) \) and \( R(\sigma_i, N) \) imply \( R(\tau_p, x \bar{M} N) \) \((1 \leq i \leq n)\).

2. \( R(\tau_p, Mx) \) and \( x \not\in \text{FV}(M) \) imply \( R(\sigma_1 \cap \ldots \cap \sigma_n \Rightarrow \tau_p, M) \), for some \( \sigma_1, \ldots, \sigma_n \).

Proof. See Lemma 8.1.

Now let us define a new computability predicate:

\[ \text{Comp}^a(\alpha, M) \iff R(\alpha, M). \]

\[ \text{Comp}^a(\sigma_1 \cap \ldots \cap \sigma_n \Rightarrow \tau_p, M) \iff (\text{Comp}^a(\sigma_i, N) \text{ implies } \text{Comp}^a(\tau_p, MN)) \]

\((1 \leq k \leq n)\).

The proof can be given following exactly the same lines than the proof of part ⇒ of Theorem 6.7, \textit{i.e.}, by proving that \( B \vdash M : \sigma_p \Rightarrow \text{Comp}^a(\sigma_p, M) \Rightarrow R(\sigma_p, M) \Rightarrow M \not\vdash_\eta. \)

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