Thunks and the $\lambda$-calculus

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Abstract

Thirty-five years ago, thunks were used to simulate call-by-name under call-by-value in Algol 60. Twenty years ago, Plotkin presented continuation-based simulations of call-by-name under call-by-value and vice versa in the $\lambda$-calculus. We connect all three of these classical simulations by factorizing the continuation-based call-by-name simulation $C_n$ with a thunk-based call-by-name simulation $T$ followed by the continuation-based call-by-value simulation $C_v$ extended to thunks.

![Diagram]

We show that $T$ actually satisfies all of Plotkin’s correctness criteria for $C_n$ (i.e., his Indifference, Simulation, and Translation theorems). Furthermore, most of the correctness theorems for $C_n$ can now be seen as simple corollaries of the corresponding theorems for $C_v$ and $T$.

Capsule Review

Many continuation-passing style (CPS) transformations are complex and can be staged into conceptually different passes. This paper shows that the call-by-name CPS transformation developed by Reynolds and Plotkin can be split into a thunk-introduction phase followed by a call-by-value CPS transformation. Moreover it proves that first phase is sufficient for the simulation purpose, formalising folklore from the days of Algol 60. The paper stands by itself, but readers may profit from having a copy of Plotkin’s 1975 paper nearby.

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1 Introduction

In his seminal paper, “Call-by-name, call-by-value and the λ-calculus” (Plotkin, 1975), Plotkin presents simulations of call-by-name by call-by-value (and vice-versa). Both of Plotkin’s simulations rely on continuations. Since Algol 60, however, programming wisdom has it that thunks can be used to obtain a simpler simulation of call-by-name by call-by-value. We show that composing a thunk-based call-by-name simulation $\mathcal{T}$ with Plotkin’s continuation-based call-by-value simulation $\mathcal{C}_v$ actually yields Plotkin’s continuation-based call-by-name simulation $\mathcal{C}_n$ (Sections 2 and 3). Revisiting Plotkin’s correctness theorems (Section 4), we provide a correction to his Translation property for $\mathcal{C}_n$, and show that the thunk-based simulation $\mathcal{T}$ satisfies all of Plotkin’s properties for $\mathcal{C}_n$. The factorization of $\mathcal{C}_n$ by $\mathcal{C}_v$ and $\mathcal{T}$ makes it possible to derive correctness properties for $\mathcal{C}_n$ from the corresponding results for $\mathcal{C}_v$ and $\mathcal{T}$. This factorization has also found several other applications already (Section 5). The extended version of this paper (Hatcliff & Danvy, 1995) gives a more detailed development as well as all proofs.

2 Continuation-based and Thunk-based Simulations

We consider $\Lambda$, the untyped $\lambda$-calculus parameterized by a set of basic constants $b$ (Plotkin, 1975, p. 127),

$$e \in \Lambda$$

$$e ::= b \mid x \mid \lambda x.e \mid e_0 \cdot e_1$$

The sets $Values_b[\Lambda]$ and $Values_x[\Lambda]$ below represent the set of values from the language $\Lambda$ under call-by-name and call-by-value evaluation respectively.

$$v \in Values_b[\Lambda] \quad \quad v \in Values_x[\Lambda]$$

$$v ::= b \mid \lambda x.e \quad \quad v ::= b \mid x \mid \lambda x.e$$

...where $e \in \Lambda$

Figure 1 displays Plotkin’s call-by-name CPS transformation $\mathcal{C}_n$ (which simulates call-by-name under call-by-value). (Side note: the term “CPS” stands for “Continuation-Passing Style”. It was coined in Steele’s MS thesis (Steele, 1978).) Figure 2 displays Plotkin’s call-by-value CPS transformation $\mathcal{C}_v$ (which simulates call-by-value under call-by-name). Figure 3 displays the standard thunk-based simulation of call-by-name using call-by-value evaluation of the language $\Lambda_r$. $\Lambda_r$ extends $\Lambda$ as follows,

$$e \in \Lambda_r$$

$$e ::= \ldots \mid delay\ e \mid force\ e$$
\( C_n(\lambda x.e) = \lambda x.C_n(\{e\}) \)

Fig. 1. Call-by-name CPS transformation

\[
\begin{align*}
C_n(\emptyset) & : \Lambda \rightarrow \Lambda \\
C_n(\{v\}) & = \lambda k. k \ C_n(\{v\}) \\
C_n(\{x\}) & = \lambda k. x \ k \\
C_n(\{e_0 \ e_1\}) & = \lambda k. C_n(\{e_0\}) (\lambda y_0. y_0 \ C_n(\{e_1\}) \ k)
\end{align*}
\]

Fig. 2. Call-by-value CPS transformation

\[
\begin{align*}
\mathcal{T} & : \Lambda \rightarrow \Lambda_r \\
\mathcal{T}(\emptyset) & = b \\
\mathcal{T}(\{x\}) & = \text{force } x \\
\mathcal{T}(\{\lambda x.e\}) & = \lambda x. \mathcal{T}(\{e\}) \\
\mathcal{T}(\{e_0 \ e_1\}) & = \mathcal{T}(\{e_0\}) \ (\text{delay } \mathcal{T}(\{e_1\}))
\end{align*}
\]

Fig. 3. Call-by-name thunk transformation

\[
\begin{align*}
C^+_n(\emptyset) & : \Lambda_r \rightarrow \Lambda \\
C^+_n(\text{force } e) & = \lambda k. C^+_n(\emptyset) (\lambda y. y \ k) \\
C^+_n(\emptyset) & : \text{Values}[\Lambda_r] \rightarrow \Lambda \\
C^+_n(\text{delay}) & = C^+_n(\emptyset)
\end{align*}
\]

Fig. 4. Call-by-value CPS transformation (extended to thunks)
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The operator delay suspends the evaluation of an expression — thereby coercing
an expression to a value. Therefore, delay e is added to the value sets in \( \Lambda_r \).

\[
v \in \text{Values}_s[\Lambda_r] \quad v \in \text{Values}_s[\Lambda_r] \quad \ldots \text{where } e \in \Lambda_r
\]

\[
v ::= \ldots | \text{delay } e \quad v ::= \ldots | \text{delay } e
\]

The operator force triggers the evaluation of such a suspended expression. This is
formalized by the following notion of reduction,

**Definition 1 (\( \tau \)-reduction)**

\[
\text{force (delay } e \text{) } \rightarrow_{\tau} e
\]

We also consider the conventional notions of reduction \( \beta, \beta_v, \eta, \) and \( \eta_v \) (Barendregt, 1984; Plotkin, 1975; Sabry & Felkeisen, 1993).

**Definition 2 (\( \beta, \beta_v, \eta, \beta_v \)-reduction)**

\[
(\lambda x. e_1) e_2 \rightarrow_{\beta} e_1[x := e_2] \\
(\lambda x. e) v \rightarrow_{\beta} e[x := v] \\
\lambda x.e \ x \rightarrow_{\eta} e \\
\lambda x.e \ x \rightarrow_{\eta_v} v
\]

\( \lambda x.e \ x \in \text{Value}_s[\Lambda] \quad x \notin \text{FV}(e) \)

For a notion of reduction \( r, \rightarrow_r \) also denotes construct compatible one-step reduction, \( \rightarrow_r \) denotes the reflexive, transitive closure of \( \rightarrow_r \), and \( =_r \) denotes the smallest equivalence relation generated by \( \rightarrow_r \) (Barendregt, 1984). We will also write \( \lambda x \vdash e_1 = e_2 \) when \( e_1 =_r e_2 \) (similarly for the other relations).

Figure 4 extends \( \mathcal{C}_\nu \) (see Figure 2) to obtain \( \mathcal{C}_\nu^+ \) which CPS-transforms thunks. \( \mathcal{C}_\nu^+ \) faithfully implements \( \tau \)-reduction in terms of \( \beta \nu \) (and thus \( \beta \) reduction). We write \( \beta \nu \) below to signify that the property holds indifferently for \( \beta_v \) and \( \beta \).

**Property 1**

For all \( e \in \Lambda_r, \lambda \beta \nu \vdash \mathcal{C}_\nu^+ \{ \text{force (delay } e \} \} = \mathcal{C}_\nu^+ \{ e \} \).  

**Proof:**

\[
\mathcal{C}_\nu^+ \{ \text{force (delay } e \} \} = \lambda k. (\lambda x k (\mathcal{C}_\nu^+ \{ e \})) (\lambda y. y k) \\
\rightarrow_{\beta_v} \lambda k. (\lambda y. y k) (\mathcal{C}_\nu^+ \{ e \}) \\
\rightarrow_{\beta} \lambda k (\mathcal{C}_\nu^+ \{ e \}) k \\
\rightarrow_{\beta} \mathcal{C}_\nu^+ \{ e \}
\]

The last step holds since a simple case analysis shows that \( \mathcal{C}_\nu^+ \{ e \} \) always has the form \( \lambda k.e' \) for some \( e' \in \Lambda \).

3 Connecting the Continuation-based and Thunk-based Simulations

\( \mathcal{C}_\nu \) can be factored into two conceptually distinct steps: (1) the suspension of argument evaluation (captured in \( T \)); and (2) the sequentialization of function application to give the usual tail-calls of CPS terms (captured in \( \mathcal{C}_\nu^+ \)).

**Theorem 1**

For all \( e \in \Lambda, \lambda \beta \nu \vdash (\mathcal{C}_\nu^+ \circ T) \{ e \} = \mathcal{C}_\nu \{ e \} \).
Proof: by induction over the structure of $e$:

case $e \equiv b$:

\[
(C^+_\eta \circ T)(b) = C^+_\eta \{b\} \\
= \lambda k. b \\
= C_\eta(b)
\]

case $e \equiv x$:

\[
(C^+_\eta \circ T)(\langle x \rangle) = C^+_\eta \{\text{force } x\} \\
= \lambda k. (\lambda k. x) (\lambda y. y k) \\
\rightarrow_{\beta_\eta} \lambda k. (\lambda y. y k) x \\
\rightarrow_{\beta_\eta} \lambda k. x k \\
= C_\eta(\langle x \rangle)
\]

case $e \equiv \lambda x. e'$:

\[
(C^+_\eta \circ T)(\langle \lambda x. e' \rangle) = C^+_\eta \{\lambda x. T(\langle e' \rangle)\} \\
= \lambda k. (\lambda x. (C^+_\eta \circ T)(\langle e' \rangle)) \\
=_{\beta_\eta} \lambda k. (\lambda x. C_\eta(\langle e' \rangle)) \quad \text{by the ind. hyp.} \\
= C_\eta(\langle \lambda x. e' \rangle)
\]

case $e \equiv e_0 \cdot e_1$:

\[
(C^+_\eta \circ T)(\langle e_0 \cdot e_1 \rangle) = C^+_\eta \{T(\langle e_0 \rangle) \cdot (T(\langle e_1 \rangle))\} \\
= \lambda k. (C^+_\eta \circ T)(\langle e_0 \rangle) (\lambda y_0. (\lambda k. (C^+_\eta \circ T)(\langle e_1 \rangle)) (\lambda y_1. y_0 y_1 k)) \\
\rightarrow_{\beta_\eta} \lambda k. (C^+_\eta \circ T)(\langle e_0 \rangle) (\lambda y_0. (\lambda y_1. y_0 y_1 k) (C^+_\eta \circ T)(\langle e_1 \rangle)) \\
\rightarrow_{\beta_\eta} \lambda k. (C^+_\eta \circ T)(\langle e_0 \rangle) (\lambda y_0. y_0 (C^+_\eta \circ T)(\langle e_1 \rangle) k) \\
=_{\beta_\eta} \lambda k. C_\eta(\langle e_0 \rangle) (\lambda y_0. y_0 C_\eta(\langle e_1 \rangle) k) \quad \text{by the ind. hyp.} \\
= C_\eta(\langle e_0 \cdot e_1 \rangle)
\]

This theorem implies that the diagram in the abstract commutes up to $\beta$-equivalence, i.e., indifferently up to $\beta_\eta$ and $\beta$-equivalence. Note that $C^+_\eta \circ T$ and $C_\eta$ only differ by administrative reductions. In fact, if we consider optimizing versions of $C_\eta$ and $C_\eta$ that remove administrative redexes, then the diagram commutes up to identity (i.e., up to $\alpha$-equivalence).

Figures 5 and 6 present two such optimizing transformations $C_{\eta, \text{opt}}$ and $C_{\eta, \text{opt}}$. The output of $C_{\eta, \text{opt}}$ is $\beta$-$\eta$ equivalent to the output of $C_\eta$, and similarly for $C_{\eta, \text{opt}}$ and $C_\eta$, as shown by Danvy and Filinski (Danvy & Filinski, 1992, pp. 378 and 387). Both are applied to the identity continuation. In Figures 5 and 6, they are presented in a two-level language à la Nielsen and Nielsen (Nielsen & Nielsen, 1992). Operationally, the overlined $\lambda$'s and $@$'s correspond to functional abstractions and applications in the program implementing the translation, while the underlined $\lambda$'s and $@$'s represent abstract-syntax constructors. It is simple to transcribe $C_{\eta, \text{opt}}$ and $C_{\eta, \text{opt}}$ into functional programs.

The optimizing transformation $C_{\eta, \text{opt}}^+$ is obtained from $C_{\eta, \text{opt}}$ by adding the following definitions:

\[
C_{\eta, \text{opt}}^+ \{\text{force } e\} \quad \text{and} \quad C_{\eta, \text{opt}}^+ \{\text{delay } e\}
\]

\[
\begin{align*}
C_{\eta, \text{opt}}^+ \{\text{force } e\} &= \lambda k. C_{\eta, \text{opt}}^+ \langle e \rangle \langle \lambda y_0. y_0 \overline{y} \langle \lambda y_1. k \overline{y} y_1 \rangle \rangle \\
C_{\eta, \text{opt}}^+ \{\text{delay } e\} &= \lambda k. C_{\eta, \text{opt}}^+ \langle e \rangle \langle \lambda y_0. k \overline{y} y_0 \rangle
\end{align*}
\]
\[ C_{\alpha, \text{opt}} \{ v, e \} = \lambda k. k \overline{\alpha (\lambda y_0, y_1, y_2. \overline{(\lambda y_1, k \overline{\alpha} y_1)) \overline{\alpha} (\lambda y_2, k \overline{\alpha} y_2))} \]

\[ C_{\alpha, \text{opt}} \{ v \} = \lambda k. k \overline{\alpha (\lambda y_0, y_1, y_2. \overline{(\lambda y_1, k \overline{\alpha} y_1)) \overline{\alpha} (\lambda y_2, k \overline{\alpha} y_2))} \]

\[ C_{\alpha, \text{opt}} \{ \lambda x, c \} = \lambda k. \lambda k. C_{\alpha, \text{opt}} \{ e \} \overline{\alpha} (\lambda y_0, \lambda y_1, \lambda y_2. \overline{(\lambda y_1, k \overline{\alpha} y_1)) \overline{\alpha} (\lambda y_2, k \overline{\alpha} y_2))} \]

\[ \text{Fig. 5. Optimizing call-by-name CPS transformation} \]

\[ C_{\nu, \text{opt}} \{ v, e \} = \lambda k. k \overline{\nu (\lambda y_0, y_1, y_2. \overline{(\lambda y_1, k \overline{\nu} y_1)) \overline{\nu} (\lambda y_2, k \overline{\nu} y_2))} \]

\[ C_{\nu, \text{opt}} \{ v \} = \lambda k. k \overline{\nu (\lambda y_0, y_1, y_2. \overline{(\lambda y_1, k \overline{\nu} y_1)) \overline{\nu} (\lambda y_2, k \overline{\nu} y_2))} \]

\[ C_{\nu, \text{opt}} \{ \lambda x, c \} = \lambda k. \lambda k. C_{\nu, \text{opt}} \{ e \} \overline{\nu} (\lambda y_0, \lambda y_1, \lambda y_2. \overline{(\lambda y_1, k \overline{\nu} y_1)) \overline{\nu} (\lambda y_2, k \overline{\nu} y_2))} \]

\[ \text{Fig. 6. Optimizing call-by-value CPS transformation} \]

Taking an operational view of these two-level specifications, the following theorem states that, for all \( e \in \Lambda \), the result of applying \( C_{\nu, \text{opt}}^+ \) to \( T \{ v \} \) (with an initial continuation \( \overline{\lambda a. a} \)) is \( \alpha \)-equivalent to the result of applying \( C_{\alpha, \text{opt}}^+ \) to \( e \) (with an initial continuation \( \overline{\lambda a. a} \)).

**Theorem 2**

For all \( e \in \Lambda \), \( (C_{\nu, \text{opt}}^+ \circ T) \{ e \} \overline{\nu} (\overline{\lambda a. a}) \equiv C_{\alpha, \text{opt}}^+ \{ e \} \overline{\alpha} (\overline{\lambda a. a}) \).

**Proof:** A simple structural induction similar to the one required in the proof of Theorem 1. We show only the case for identifiers (the others are similar). The overlined constructs are computed at translation time, and thus simplifying overlined constructs using \( \beta \)-conversion yields equivalent specifications.

**case** \( e \equiv x \): \( (C_{\nu, \text{opt}}^+ \circ T) \{ x \} = \lambda k. (\overline{\lambda k. k \overline{\nu} x} \overline{\nu} (\lambda y_0, y_1. \overline{(\lambda y_1, k \overline{\nu} y_1)) \overline{\nu} (\lambda y_2, k \overline{\nu} y_2))} \)

\[ = \lambda k. (\lambda y_0, y_1. \overline{(\lambda y_1, k \overline{\nu} y_1)) \overline{\nu} x} \]

\[ = \lambda k. x \overline{\nu} (\lambda y_1, k \overline{\nu} y_1) \]

\[ = C_{\alpha, \text{opt}}^+ \{ x \} \]
4 Revisiting Plotkin’s Correctness Properties

Figure 7 presents single-step evaluation rules specifying the call-by-name and call-by-value operational semantics of $\Lambda$ programs (closed terms). The (partial) evaluation functions $\text{eval}_n$ and $\text{eval}_v$ are defined in terms of the reflexive, transitive closure (denoted $\rightarrow^*$) of the single-step evaluation rules.

$$\text{eval}_n(e) = v \iff e \rightarrow^*_n v$$
$$\text{eval}_v(e) = v \iff e \rightarrow^*_v v$$

The evaluation rules for $\Lambda_r$ are obtained by adding the following rules to both the call-by-name and call-by-value evaluation rules of Figure 7.

$$e \rightarrow e'$$
$$\text{force} e \rightarrow \text{force} e'$$

For a language $l$, $\text{Programs}[l]$ denotes the closed terms in $l$. For meta-language expressions $E_1, E_2$, we write $E_1 \equiv E_2$ when $E_1$ and $E_2$ are both undefined, or else both are defined and denote $\alpha$-equivalent terms. We will also write $E_1 \equiv_r E_2$ when $E_1$ and $E_2$ are both undefined, or else are both defined and denote $r$-convertible terms for the convertibility relation generated by some notion of reduction $r$.

Plotkin expressed the correctness of his simulations $\mathcal{C}_0$ and $\mathcal{C}_v$ via three properties: **Indifference**, **Simulation**, and **Translation**. **Indifference** states that call-by-name and call-by-value evaluation coincide on terms in the image of the CPS transformation. **Simulation** states that the desired evaluation strategy is properly simulated. **Translation** states how the transformation relates program calculi for each evaluation strategy (e.g., $\lambda\beta$, $\lambda\beta\xi$). Let us restate these properties for Plotkin’s original presentation of $\mathcal{C}_0$ (hereby noted $P_0$) (Plotkin, 1975, p. 153), that only differs from Figure 1 at the line for identifiers.

$$P_0(x) = x$$
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Theorem 3 (Plotkin 1975)

For all $e \in \text{Programs}[\Lambda],$

1. **Indifference:** $\text{eval}_e(\mathcal{P}_n[e]_I) \simeq \text{eval}_e(\mathcal{P}_n[e]_I)$

2. **Simulation:** $\mathcal{P}_n(\text{eval}_e(e)) \simeq \text{eval}_e(\mathcal{P}_n[e]_I)$

where $I$ denotes the identity function and is used as the initial continuation.

Plotkin also claimed the following **Translation** property.

Claim 1 (Plotkin 1975)

For all $e_1, e_2 \in \Lambda,$

**Translation:** $\lambda\beta \vdash e_1 = e_2 \iff \lambda\beta \vdash \mathcal{P}_n[e_1] = \mathcal{P}_n[e_2]$

where $\iff$ is the equivalence between two terms.

The *Translation* property purports to show that $\beta$-equivalence classes are preserved and reflected by $\mathcal{P}_n.$ The property, however, does not hold because

$$\lambda\beta \vdash e_1 = e_2 \not\iff \lambda\beta \vdash \mathcal{P}_n[e_1] = \mathcal{P}_n[e_2].$$

The proof breaks down at the statement “It is straightforward to show that $\lambda\beta \vdash e_1 = e_2$ implies $\lambda\beta \vdash \mathcal{P}_n[e_1] = \mathcal{P}_n[e_2]$ ...” (Plotkin, 1975, p. 158). In some cases, $\eta_v$ is needed to establish the equivalence of the CPS-images of two $\beta$-convertible terms. For example, $\lambda x. (\lambda z. z) x \rightarrow_{\beta} \lambda x. x$ but

$$\mathcal{P}_n[\lambda x. (\lambda z. z) x] = \lambda k. k (\lambda x. \lambda k. (\lambda k. k (\lambda z. z)) (\lambda y. y x k)) \quad (1)$$

$$\rightarrow_{\beta} \lambda k. k (\lambda x. \lambda y. \lambda x. y x k) (\lambda z. z) \quad (2)$$

$$\rightarrow_{\beta} \lambda k. k (\lambda x. \lambda y. x k) \quad (3)$$

$$\rightarrow_{\beta} \lambda k. k (\lambda x. x k) \quad (4)$$

$$\rightarrow_{\eta_v} \lambda k. k (\lambda x. x) \quad \ldots \eta_v \text{ is needed for this step} \quad (5)$$

$$\mathcal{P}_n[\lambda x. x] = \mathcal{P}_n[\lambda x. (\lambda z. z) x] \quad (6)$$

Since the two distinct terms at lines (4) and (5) are $\beta$-normal, confluence of $\beta$ implies $\lambda\beta \not\vdash \mathcal{P}_n[e_1] = \mathcal{P}_n[e_2].$

In practice, though, $\eta_v$ reductions such as those required in the example above are unproblematic if they are embedded in proper CPS contexts (e.g., contexts in the language of terms in the image of $\mathcal{P}_n$ closed under $\beta$ reductions). When $\lambda k (\lambda x. \lambda k. x k)$ is embedded in a CPS context, $x$ will always bind to a term of the form $\lambda k. c$ during evaluation. In this case, the $\eta_v$ reduction can be expressed by a $\beta_v$ reduction. If the term, however, is not embedded in a CPS context (e.g., $\lambda y. [\lambda y, y b]$), the $\eta_v$ reduction is unsound, i.e., it fails to preserve operational equivalence as defined by Plotkin (Plotkin, 1975, pp. 144,147). Such reductions are unsound due to “improper” uses of basic constants. For example, $\lambda x. b x \rightarrow_{\eta_v} b$ but $\lambda x. b x \not\rightarrow_{\beta_v} b$ (take $C = [\lambda]$ where $\approx_v$ is the call-by-value operational equivalence relation defined
by Plotkin (Hatcliff & Danvy, 1995, p. 9). Note, finally, that a simple typing discipline eliminates improper uses of basic constants, and consequently gives soundness for $\eta$. 

The simplest solution for recovering the Translation property is to change the translation of identifiers from $P_n\{x\} = x$ to $\lambda k. x k$ — obtaining the translation $C_n$ given in Figure 1.†

For the example above, the modified translation gives

$$\lambda x. (\lambda z. z) x \vdash C_n(\lambda x. (\lambda z. z) x) = C_n(\lambda x. x).$$

The following theorem gives the correctness properties for $C_n$.

Theorem 4

For all $c \in \text{Programs}[A]$ and $c_1, c_2 \in A$,

1. Indifference: $\text{eval}_c(C_n(\{c\} I) \simeq_\beta \text{eval}_c(C_n(\{c\} I))$

2. Simulation: $C_n(\text{eval}_c(c)) \simeq_\beta C_n(\{c\} I)$

3. Translation: $\lambda \beta \vdash c_1 = c_2$ \iff $\lambda \beta \vdash C_n(\{c_1\}) = C_n(\{c_2\})$
   \hspace{1cm} \iff $\lambda \beta \vdash C_n(\{c_1\} I) = C_n(\{c_2\} I)$
   \hspace{1cm} \iff $\lambda \beta \vdash C_n(\{c_1\} I) = C_n(\{c_2\} I)$

The Indifference and Translation properties remain the same. The Simulation property, however, holds up to $\beta$-equivalence while Plotkin’s Simulation for $P_n$ holds up to $\alpha$-equivalence. For example,

$$C_n(\text{eval}_n((\lambda z. \lambda y. z) b)) = \lambda y. \lambda k. k b$$

whereas

$$\text{eval}_n(C_n(\{\lambda z. \lambda y. z\} b)) = \lambda y. \lambda k. (\lambda k. k b) k.$$

In fact, proofs of Indifference, Simulation, and most of the Translation can be derived from the correctness properties of $C_n^+$ and $T$ (see Section 5). All that remains of Translation is to show that $\lambda \beta \vdash C_n(\{c_1\} I) = C_n(\{c_2\} I)$ implies $\lambda \beta \vdash c_1 = c_2$ and this follows in a straightforward manner from Plotkin’s original proof for $P_n$ (Hatcliff & Danvy, 1995, p. 31). The following theorem gives the Indifference, Simulation, and Translation properties for $C_n$.

Theorem 5 (Plotkin 1975)

For all $c \in \text{Programs}[A]$ and $c_1, c_2 \in A$,

1. Indifference: $\text{eval}_c(C_n(\{c\} I) \simeq_\beta \text{eval}_c(C_n(\{c\} I))$

2. Simulation: $C_n(\text{eval}_c(c)) \simeq_\beta C_n(\{c\} I)$

3. Translation: $\lambda \beta \vdash c_1 = c_2$ then $\lambda \beta \vdash C_n(\{c_1\}) = C_n(\{c_2\})$
   \hspace{1cm} Also $\lambda \beta \vdash C_n(\{c_1\} I) = C_n(\{c_2\} I)$ iff $\lambda \beta \vdash C_n(\{c_1\} I) = C_n(\{c_2\} I)$

† In the context of Parigot’s $\lambda \mu$-calculus (Parigot, 1992), de Groot independently noted the problem with Plotkin’s Translation theorem and proposed a similar correction (de Groot, 1994).
The **Translation** property states that β-equivalent terms are also convertible in the image of $\mathcal{C}_v$. In contrast to the theory $\lambda\beta$ appearing in the **Translation** property for $\mathcal{C}_n$ (Theorem 4), the theory $\lambda\beta_v$ is **incomplete** in the sense that it cannot prove the equivalence of some terms whose CPS images are provably equivalent using $\lambda\beta$ or $\lambda\beta_v$ (Sabry & Felkiesen, 1993). The properties of $\mathcal{C}_v$ as stated in Theorem 5 can be extended to the transformation $\mathcal{C}_v^+$ defined on the language $T$ — the set of terms in the image of $T$ closed under $\beta_v\tau$ reduction. It is straightforward to show that the following grammar generates exactly the set of terms $T$ (Hatcliff & Danvy, 1995, pp. 32,33).

\[
\text{t ::= b | force x | force (delay t) | } \lambda x . t | t_0 (\text{delay } t_1)
\]

**Theorem 6**

For all $t \in \text{Programs}[T]$ and $t_1, t_2 \in T$,

1. **Indifference:** $\text{eval}_b(\mathcal{C}_v^+ \langle t \rangle L) \simeq \text{eval}_b(\mathcal{C}_v^+ \langle t \rangle L)$
2. **Simulation:** $\mathcal{C}_v^+ (\text{eval}_b(t)) \simeq \text{eval}_b(\mathcal{C}_v^+ \langle t \rangle L)$
3. **Translation:** If $\lambda\beta_v \tau \vdash t_1 = t_2$ then $\lambda\beta_v \vdash \mathcal{C}_v^+ \langle t_1 \rangle = \mathcal{C}_v^+ \langle t_2 \rangle$

   Also $\lambda\beta_v \vdash \mathcal{C}_v^+ \langle t_1 \rangle = \mathcal{C}_v^+ \langle t_2 \rangle$ if $\lambda\beta_v \vdash \mathcal{C}_v^+ \langle t_1 \rangle = \mathcal{C}_v^+ \langle t_2 \rangle$

**Proof:** For **Indifference** and **Simulation** it is only necessary to extend Plotkin’s colon-translation proof technique and definition of stuck terms to account for delay and force. The proofs then proceed along the same lines as Plotkin’s original proofs for $\mathcal{C}_v$ (Plotkin, 1975, pp. 148-152). **Translation** follows from the **Translation** component of Theorem 5 and Property 1 (Hatcliff & Danvy, 1995, p. 39).

Thunks are sufficient for establishing a call-by-name simulation satisfying all of the correctness properties of the continuation-passing simulation $\mathcal{C}_n$. Specifically, we prove the following theorem which recasts the correctness theorem for $\mathcal{C}_n$ (Theorem 4) in terms of $T$. The last two assertions of the **Translation** component of Theorem 4 do not appear here since the identity function as the initial continuation only plays a role in CPS evaluation.

**Theorem 7**

For all $c \in \text{Programs}[\Lambda]$ and $c_1, c_2 \in \Lambda$,

1. **Indifference:** $\text{eval}_b(\langle T \langle c \rangle \rangle) \simeq \text{eval}_b(\langle T \langle c \rangle \rangle)$
2. **Simulation:** $T \langle \text{eval}_b(c) \rangle \simeq \text{eval}_b(\langle T \langle c \rangle \rangle)$
3. **Translation:** $\lambda\beta_\tau \vdash c_1 = c_2$ if $\lambda\beta_\tau \vdash \langle T \langle c_1 \rangle \rangle = \langle T \langle c_2 \rangle \rangle$

   if $\lambda\beta_\tau \vdash \langle T \langle c_1 \rangle \rangle = \langle T \langle c_2 \rangle \rangle$

**Proof:** The proof of **Indifference** is trivial: one can intuitively see from the grammar for $T$ (which includes the set of terms in the image of $T$ closed under evaluation steps) that call-by-name and call-by-value evaluation will coincide since all function arguments are values.

The proof of **Simulation** is somewhat involved. It begins by inductively defining a relation $\tilde{\sim} \subseteq \Lambda \times \Lambda_\tau$ such that $c \tilde{\sim} t$ holds exactly when $\lambda\tau \vdash \langle T \langle c \rangle \rangle = t$. The crucial step is then to show that for all $c \in \text{Programs}[\Lambda]$ and $t \in \text{Programs}[\Lambda_\tau]$
\[ T_c : \Lambda \rightarrow \Lambda \]
\[ T_c(\beta) = b \]
\[ T_c(x) = x \ b \quad \text{for some arbitrary basic constant } b \]
\[ T_c(\lambda x.e) = \lambda x. T_c(e) \]
\[ T_c(e_1 e_2) = T_c(\eta)(\lambda z. T_c(e_1)) \quad \text{where } z \not\in \text{FV}(e_1) \]

Fig. 8. Thunk introduction implemented in \( \Lambda \)

such that \( e \sim t, e \xrightarrow{n} e' \) implies that there exists a \( t' \) such that \( t \xrightarrow{n} t' \) and \( e' \sim t' \) (Hatcliff & Danvy, 1995, Sect. 2.3.2).

**Translation** is established by first defining a translation \( \mathcal{T}^{-1} : \Lambda_r \rightarrow \Lambda \) that simply removes delay and force constructs. One then shows that \( \mathcal{T} \) and \( \mathcal{T}^{-1} \) establish an equational correspondence (Sabry & Felleisen, 1993) (or more precisely a reflection (Sabry & Wandler, 1996)) between theories \( \lambda \beta \) and \( \lambda \beta \tau \) (and \( \lambda \beta \) and \( \lambda \beta \tau \)). **Translation** follows as a corollary of this stronger result (Hatcliff & Danvy, 1995, Sect. 2.3.3).

Representing thunks via abstract suspension operators delay and force simplifies the technical presentation and enables the connection between \( \mathcal{C}_n \) and \( \mathcal{C}_v \) presented in Section 3. Elsewhere (Hatcliff, 1994), we show that the delay/force representation of thunks and associated properties (i.e., reduction properties and translation into CPS) are not arbitrary, but are determined by the relationship between strictness and continuation monads (Moggi, 1991).

Figure 8 presents the transformation \( T_c \) that implements thunks directly in \( \Lambda \) using what Plotkin described as the “protecting by a \( \lambda \) technique” (Plotkin, 1975, p. 147). An expression is delayed by wrapping it in an abstraction with a dummy parameter. A thunk is forced by applying it to a dummy argument.

The following theorem recasts the correctness theorem for \( \mathcal{C}_n \) (Theorem 4) in terms of \( T_c \).

**Theorem 8**
For all \( e \in \text{Programs}[\Lambda] \) and \( e_1, e_2 \in \Lambda \),

1. **Indifference:** \( \text{eval}_n(T_c(e)) \simeq \text{eval}_n(T_c(e')) \)
2. **Simulation:** \( T_c(\text{eval}_n(e)) \simeq_\beta_n \text{eval}_n(T_c(e)) \)
3. **Translation:** \( \lambda \beta \vdash e_1 = e_2 \iff \lambda \beta \vdash T_c(e_1) = T_c(e_2) \)

**Proof:** Follows the same pattern as the proof of Theorem 7 (Hatcliff & Danvy, 1995, Sect. 2.4).

5 **Applications**

5.1 **Deriving correctness properties of \( \mathcal{C}_n \)**

When working with CPS, one often needs to establish technical properties for both a call-by-name and a call-by-value CPS transformation. This requires two sets of
Thunks and the \( \lambda \)-calculus

Proofs that both involve CPS. By appealing to the factoring property, however, often only one set of proofs over call-by-value CPS terms is necessary. The second set of proofs deals with thunked terms which have a simpler structure. For instance, Indifference and Simulation for \( \mathcal{C}_n \) follow from Indifference and Simulation for \( \mathcal{C}^+_n \) and \( \mathcal{T} \) and Theorem 1. Here we show only the results where evaluation is undefined or results in a basic constant \( b \). See (Hatchiff & Danvy, 1995, p. 31) for a derivation of \( \mathcal{C}_n \) Simulation for arbitrary results.

For Indifference, let \( e, b \in \Lambda \) where \( b \) is a basic constant. Then

\[
\text{eval}_n (\mathcal{C}_n [e] (\lambda y. y)) = b
\]

\( \Leftrightarrow \)

\[
\text{eval}_n((\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y)) = b \quad \text{...Theorem 1 and the soundness of } \beta_n
\]

\[
\text{eval}_n((\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y)) = b \quad \text{...Theorem 6 (Indifference)}
\]

\[
\text{eval}_n(\mathcal{C}_n [e] (\lambda y. y)) = b
\]

...Theorem 1 and the soundness of \( \beta \)

For Simulation, let \( e, b \in \Lambda \) where \( b \) is a basic constant, Then

\[
\text{eval}_n (e) = b
\]

\( \Leftrightarrow \)

\[
\text{eval}_n(\mathcal{T}(\lambda y. y)) = b \quad \text{...Theorem 7 (Simulation)}
\]

\[
\text{eval}_n((\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y)) = b \quad \text{...Theorem 6 (Simulation)}
\]

\[
\text{eval}_n((\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y)) = b \quad \text{...Theorem 6 (Indifference)}
\]

\[
\text{eval}_n(\mathcal{C}_n [e] (\lambda y. y)) = b
\]

...Theorem 1 and the soundness of \( \beta \)

For Translation, it is not possible to establish Theorem 4 (Translation for \( \mathcal{C}_n \)) in the manner above since Theorem 6 (Translation for \( \mathcal{C}^+_n \)) is weaker in comparison. However, the following weaker version can be derived. Let \( e_1, e_2 \in \Lambda \). Then

\[
\lambda \beta \vdash e_1 = e_2
\]

\( \Leftrightarrow \)

\[
\lambda \beta \vdash \mathcal{T}(\lambda y. y) = \mathcal{T}(\lambda y. y)
\]

...Theorem 7 (Translation)

\[
\lambda \beta \vdash (\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y) = (\mathcal{C}_n^+ \circ \mathcal{T})(\lambda y. y)
\]

...Theorem 6 (Translation)

\( \Rightarrow \)

\[
\lambda \beta \vdash \mathcal{C}_n [e_1] = \mathcal{C}_n [e_2]
\]

...Theorem 1

\[
\lambda \beta \vdash \mathcal{C}_n [e_1] I = \mathcal{C}_n [e_2] I
\]

...compatibility of \( =_\beta \).

5.2 Deriving a CPS transformation directed by strictness information

Strictness information indicates arguments that may be safely evaluated eagerly (i.e., without being delayed) — in effect, reducing the number of thunks needed in a program and the overhead associated with creating and evaluating suspensions (Bloss et al., 1988; Mycroft, 1981; Okasaki et al., 1994). In an earlier work (Danvy & Hatchiff, 1993), we gave a transformation \( \mathcal{T} \) that optimizes thunk introduction based on strictness information. We then used the factorization of \( \mathcal{C}_n \) by \( \mathcal{C}^+_n \) and \( \mathcal{T} \) to derive an optimized CPS transformation \( \mathcal{C}_s \) for strictness-analyzed call-by-name terms. This staged approach can be contrasted with Burn and Le Métayer’s monolithic strategy (Burn & Le Métayer, 1996).

The resulting transformation \( \mathcal{C}_s \) yields both call-by-name-like and call-by-value-like continuation-passing terms. Due to the factorization, the proof of correctness for the optimized transformation follows as a corollary of the correctness of the strictness analysis and the correctness of \( \mathcal{T} \) and \( \mathcal{C}^+_n \).

Amtsoft (Amtsoft, 1993) and Steckler and Wand (Steckler & Wand, 1994) have
proven the correctness of transformations which optimize the introduction of thunks based on strictness information.

5.3 Deriving a call-by-need CPS transformation

Okasaki, Lee, and Tarliti (Okasaki et al., 1994) have also applied the factorization to obtain a "call-by-need CPS transformation" $C_{\text{need}}$. The lazy evaluation strategy characterizing call-by-need is captured with memo-thunks (Blass et al., 1988). $C_{\text{need}}$ is obtained by extending $C_{\nu}$ to transform memo-thunks to CPS terms with store operations (which are used to implement the memoization) and composing it with the memo-thunk introduction.

Okasaki et al. optimize $C_{\text{need}}$ by using strictness information along the lines discussed above. They also use sharing information to detect where memo-thunks can be replaced by ordinary thunks. In both cases, optimizations are achieved by working with simpler thunked terms as opposed to working directly with CPS terms.

5.4 Alternative CPS transformations

Thunks can be used to factor a variety of call-by-name CPS transformations. In addition to those discussed here, one can factor a variant of Reynolds’s CPS transformation directed by strictness information (Hatchfield, 1994; Reynolds, 1974), as well as a call-by-name analogue of Fischer’s call-by-value CPS transformation (Fischer, 1993; Sabry & Felkeisen, 1993).

Obtaining the desired call-by-name CPS transformation via $C_{\nu}$ and $T$ depends on the representation of thunks. For example, if one works with $T_L$ (see Figure 8) instead of $T$, $C_{\nu} \circ T_L$ still gives a valid CPS simulation of call-by-name by call-value. However, $\beta$ equivalence with $C_{\nu}$ is not obtained (i.e., $\lambda x. x y \not\equiv C_{\nu} \circ T_L (\epsilon x y)$), as shown by the following derivations.

$$(C_{\nu} \circ T_L) \epsilon x = C_{\nu} \{x b\}$$

$$= \lambda k.(x b) k$$

$$(C_{\nu} \circ T_L) \epsilon_0 \epsilon_1 = C_{\nu} \{T_L \{\epsilon_0\} (\lambda z . T_L \{\epsilon_1\})\}$$

$$= \lambda k.(C_{\nu} \circ T_L) \epsilon_0 (\lambda g.(g \lambda z . (C_{\nu} \circ T_L) \epsilon_1)) k$$

The representation of thunks given by $T_L$ is too concrete in the sense that the delaying and forcing of computation is achieved using specific instances of the more general abstraction and application constructs. When composed with $T_L$, $C_{\nu}$ treats the specific instances of thunks in their full generality, and the resulting CPS terms contain a level of inessential encoding of delay and force.

5.5 The factorization holds for types

Plotkin's continuation-passing transformations were originally stated in terms of untyped $\lambda$-calculus. These transformations have been shown to preserve well-typedness
of terms (Griffin, 1990; Harper & Lillibridge, 1993; Meyer & Wand, 1985; Murthy, 1990). The thunk transformation \( T \) also preserves well-typedness of terms, and the relationship between \( \mathcal{C}^\sharp \circ T \) and \( \mathcal{C}_n \) is reflected in transformations on types (Hatcliff & Danvy, 1995, Sect. 4).

6 Related Work

Ingerman (Ingerman, 1961), in his work on the implementation of Algol 60, gave a general technique for generating machine code implementing procedure parameter passing. The term thunk was coined to refer to the compiled representation of a delayed expression as it gets pushed on the control stack (Raymond (editor), 1992). Since then, the term thunk has been applied to other higher-level representations of delayed expressions and we have followed this practice.

Bloss, Hudak, and Young (Bloss et al., 1988) study thunks as the basis of an implementation of lazy evaluation, Optimizations associated with lazy evaluation (e.g., overwriting a forced expression with its resulting value) are encapsulated in the thunk. They give several representations with differing effects on space and time overhead.

Riecke (Riecke, 1991) has used thunks to obtain fully abstract translations between versions of PCF with differing evaluation strategies. In effect, he establishes a fully abstract version of the Simulation property for thunks. The Indifference property is also immediate for Riecke since all function arguments are values in the image of his translation (and this property is maintained under reductions). The thunk translation required for full abstraction is much more complicated than our transformation \( T \) and consequently it cannot be used to factor \( \mathcal{C}_n \). In addition, since Riecke’s translation is based on typed-indexed retractions, it does not seem possible to use it (and the corresponding results) in an untyped setting as we require here.

Asperti and Curien formulate thunks in a categorical setting (Asperti, 1992; Curien, 1986). Two combinators freeze and unfreeze, which are analogous to delay and force but have slightly different equational properties, are used to implement lazy evaluation in the Categorical Abstract Machine. In addition, freeze and unfreeze can be elegantly characterized using a co-monad.

In his original paper (Plotkin, 1975, p. 147), Plotkin acknowledges that thunks provide some simulation properties but states that “...these protecting by a \( \lambda \) techniques do not seem to be extendable to a complete simulation and it is fortunate that the technique of continuations is available.” (Plotkin, 1975, p. 147). By “protecting by a \( \lambda \)”, Plotkin refers to a representation of thunks as \( \lambda \)-abstractions with a dummy parameter, as in Figure 8. In a set of unpublished notes, however, he later showed that the “protecting by a \( \lambda \)” technique is sufficient for a complete simulation (Plotkin, 1978).

An earlier version of Section 3 appeared in the proceedings of WSA’92 (Danvy & Hatcliff, 1992). Most of these proofs have been checked in Elf (Pfenning, 1991) by Niss and the first author (Niss & Hatcliff, 1995). Elsewhere (Hatcliff, 1994), we also consider an optimizing version of \( T \) that does not introduce thunks for identifiers.
occurring as function arguments:

\[ \mathcal{T}_{opt}(\ell x) = \mathcal{T}_{opt}(v) x \]

\( \mathcal{T}_{opt} \) generates a language \( \mathcal{T}_{opt} \) which is more refined than \( \mathcal{T} \) (referred to in Theorem 6).

Finally, Lawall and Danvy investigate staging the call-by-value CPS transformation into conceptually different passes elsewhere (Lawall & Danvy, 1993).

7 Conclusion

We have connected the traditional thunk-based simulation \( \mathcal{T} \) of call-by-name under call-by-value and Plotkin’s continuation-based simulations \( \mathcal{C}_o \) and \( \mathcal{C}_v \) of call-by-name and call-by-value. Almost all of the technical properties Plotkin established for \( \mathcal{C}_o \) follow from the properties of \( \mathcal{T} \) and \( \mathcal{C}_v^+ \) (the extension of \( \mathcal{C}_v \) to thunks). When reasoning about \( \mathcal{C}_o \) and \( \mathcal{C}_v \), it is thus often sufficient to reason about \( \mathcal{C}_v^+ \) and the simpler simulation \( \mathcal{T} \). We have also given several applications involving deriving optimized continuation-based simulations for call-by-name and call-by-need languages and performing CPS transformation after static program analysis.

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