TOWARDS A SEMANTIC THEORY of
DYNAMIC BINDING,

by

Michael Gordon

Research sponsored by
Advanced Research Projects Agency
ARPA Order No. 2494

COMPUTER SCIENCE DEPARTMENT
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TOWARDS A SEMANTIC THEORY OF DYNAMIC BINDING

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Abstract

The results in this paper contribute to the formulation of a semantic theory of dynamic binding (fluid variables). The axioms and theorems are language independent in that they don't talk about programs - i.e. syntactic objects - but just about elements in certain domains. Firstly the equivalence (in the circumstances where it's true) of "tying a knot" through the environment (elaborated in the paper) and taking a least fixed point is shown. This is central in proving the correctness of LISP "eval" type interpreters. Secondly the relation which must hold between two environments if a program is to have the same meaning in both is established. It is shown how the theory can be applied to LISP to yield previously known facts.
ACKNOWLEDGEMENTS

Thanks to John Allen, Rod Burstall, Friedrich von Henke, Robert Milne, Gordon Plotkin, Bob Tennent and Chris Wadsworth for helpful discussions and correspondence. John Allen, Dane Scott and Akinori Yonezawa suggested improvements and pointed out errors in preliminary drafts of this report.

This research was supported in part by the Advanced Research Projects Agency of the Office of the Secretary of Defense under contract DAHC 15-73-C-0435, ARPA order no. 2494.

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1. Introduction

The art of semantics is now sufficiently developed that most computer languages can be given concise, elegant and intuitive formal descriptions. The theory of these descriptions is well enough understood that useful facts - such as the correctness of implementations - are fairly straightforward to prove. Unfortunately proofs tend to be very long and the results obtained rather lacking in generality. For example the proof of correctness of an implementation for one language has to be redone for a similar implementation of another. Of course once the proof idea is known no real creative acts are needed in applying it and thus a certain amount of generality is obtained. However this generality isn't of a type that's easy to use (except, perhaps, by people with considerable knowledge of the underlying theory). A more direct way of being general is to isolate explicitly the assumptions used and then to prove the results from these. Then to apply such a result one just needs to check the language satisfies the appropriate "axioms" - and this will normally be much less demanding than redoing a whole proof by analogy with an existing one.

In this note I've formulated abstract versions of two results about languages which use dynamic binding of free variables. Initially these were proved for LISP (they were needed in proving the correctness of an implementation). The abstract versions described below can be instantiated to yield the LISP ones. Although the two results proven are completely language-independent (in that they don't talk about programs - i.e. syntactic objects - but just about elements in certain domains) they aren't as general as one might hope. Some situations in which dynamic binding is used and which intuitively should fall under their compass don't. This is a defect of the present work - I don't think it's a necessary difficulty.
2. Informal Discussion of Results

When reasoning about programs it’s often useful to be able to exhibit the denotation of a recursive procedure as the least fixed point of some functional. Doing this enables, for example, computation-induction to be used. The first result to be discussed helps with this as it concerns the equivalence (in certain circumstances) of "tying a knot" through the environment (elaborated below) with taking a least fixed point. Besides being of interest in its own right, this result is at the heart of the correctness of LISP eval type interpreters. Hopefully the abstract version will assist in proving the correctness of similar interpreters for other languages.

The way recursive definitions are handled by many LISP implementations is to bind the body of the function to its own name on the alist. This creates a circularity or "knot" in which places inside the function body (namely recursive calls) point back to the beginning of the function. Now the standard analysis of recursion is via the Y-operator (i.e. in terms of least fixed points) and consequently in proving the correctness of "knotting" interpreters with respect to standard semantics it’s necessary to ascertain the conditions under which "knotting" and fixedpointing are equivalent. Contrary to what one might expect they aren’t always the same. This is shown below.

The second result concerns what relation needs to hold between two environments a,a' (alists in the case of LISP) for a form e to evaluate to the same values in both a and a'. Call this condition "a≈a'".
A first guess might be that the two environments must agree on the free variables of \( e \) (as is the case for terms in predicate-calculus or the \( \lambda \)-calculus). This won't do however for although \( a \) and \( a' \) might agree on \( e \)'s free variables the things they bind to these might depend on other variables not free in \( e \) and on which \( a \) and \( a' \) differ (e.g. if \( e=x \), \( a \) and \( a' \) both bind \( x \) to \( y \) but \( a \) binds \( y \) to \( 1 \) whilst \( a' \) binds it to \( 2 \)). What is clearly needed is that \( a \) and \( a' \) agree on \( e \)'s free variables and on the variables free in the things bound to these variables .... etc.

To formulate this for LISP one just needs a recursive definition like:

\[
a = a' \iff \forall x. [ x \text{ free in } e \Rightarrow a(x) = a'(x) \text{ and } a = F(a) = a' ]
\]

Now given a syntax for \( e \)'s it's easy to formalise "\( x \) free in \( e \)" - the difficulty arises if one wants a syntax independent definition. What's needed is an abstract notion of free-ness applicable to elements of the type denoted by \( e \) (and hopefully denoted also by programs from languages other than LISP). I describe such a notion below.

### 3. Formalization

#### 3.1. Knots and Fixed-points

Before proceeding with abstract formulations of the above it's necessary to describe the environments needed to handle dynamic binding. Let \( D \) be an arbitrary domain of expression values and let \( Env = Id \rightarrow V_0 \) be the associated domain of environments. Elements of \( V_0 \) are - in the case of dynamic binding - denotations of objects which may contain free variables and so might still depend on the environment. Hence \( V_0 = Env \rightarrow D \) and thus \( Env \) must satisfy \( Env = Id \rightarrow (Env \rightarrow D) \).
It turns out to be necessary (see lemma 8 below) to require in addition that if \( \rho \in Env \) then \( \rho \) is strict i.e. \( \rho(\bot) = \bot \) thus if \( /D_1 \rightarrow D_2 / \) is the domain of strict continuous functions from \( D_1 \) to \( D_2 \) then \( Env \) must have type satisfying:

\[
Env = Id \rightarrow (Env \rightarrow D).
\]

From this one can immediately formulate what it means for "knotting" and fixedpointing to be the same viz. we require for \( \nu \in \nu_0 \) and \( \rho \in Env \):

\[
\nu(\rho[v/x]) = Y(F_x(v))\rho \quad \text{where } F_x(v) = \nu v'. \rho p'. v(p'[v'/x])
\]

\[
\text{knot} \quad \text{fixedpoint}
\]

Here \( \rho[v/x] \) is \( \rho \) updated to bind \( v \) to \( x \). Unfortunately this equality isn't true in general.

For example if:

\[
\begin{align*}
v &= \lambda p'. p'(y)p' \\
\rho &= \bot[(\lambda p'. d)/x]([\lambda p'. p'(x)p])/y)
\end{align*}
\]

Then it turns out that \( \nu(\rho[v/x]) = d \neq \bot = Y(F_x(v))\rho \).

For we have:

\[
\begin{align*}
\nu(\rho[v/x]) &= (\rho[v/x])(y)(\rho[v/x]) \\
&= \nu(y)(\rho[v/x]) \\
&= (\lambda p'. p'(x)p)(\rho[v/x]) \\
&= (\rho[v/x])(x)p \\
&= v(p) \\
&= \rho(y)p \\
&= (\lambda p'. p'(x)p)\rho \\
&= \rho(x)p \\
&= (\lambda p'. d)p \\
&= d
\end{align*}
\]

And as \( Y(F_x(v))\rho = \cup \rho F_x(v)(\bot)\rho \) and
\[ F_x(v)^{(\bot)} p = \bot \] implies

\[
F_x(v)^{(\bot)} p = F_x(v)(F_x(v)^{(\bot)}) p = v(p[F_x(v)^{(\bot)}/x]) = (\rho[F_x(v)^{(\bot)}/x])(y)(\rho[F_x(v)^{(\bot)}/x]) \quad \text{(by definition of } v) \\
= (\lambda y. \rho^*(x) p \rho[F_x(v)^{(\bot)}/x]) \quad \text{(by definition of } \rho) \\
= F_x(v)^{(\bot)} p = \bot
\]

It follows by induction on \( n \) that: \( \forall n. F_x(v)^{(\bot)} p = \bot \) and so \( Y(F_x(v))^p = \bot \).

In [1] and [2] it is shown that for \( v \)'s and \( p \)'s which are the denotations of LISP functions and \( \alpha \)lists respectively the equation \( v(p[v/x]) = Y(F_x(v))^p \) does in fact hold. The proof used was very specific to LISP (being essentially an induction on the size of computations on a certain abstract interpreter). Now hopefully the result should hold for dynamic binding in general rather than just for LISP. Thus the problem arises of isolating and stating those properties of dynamic binding which, when possessed by \( v \) and \( p \), entail \( v(p[v/x]) = Y(F_x(v))^p \). To do this we need to introduce recursively defined (but not necessarily monotonic) relations of the type first studied by Milne [5] and Reynolds [7]. Using these we can then provide a (partial) abstract characterisation of dynamic binding by defining a notion of "regular" for which:

\[
\forall \rho. \rho \text{ regular } \Rightarrow v(p[v/x]) = Y(F_x(v))^p
\]

From now on \( x, x', x'', \ldots, y, y', y'' \) etc. will range over \( \text{Id} \). \( X, Y \) will range over subsets of \( \text{Id} \). \( \rho, \rho', \rho'' \) will range over \( \text{Env} \). \( \nu, \nu', \nu'' \) will range over \( \nu_0 = \text{Env} \rightarrow D \) and \( d, d', d'' \) will range over \( D \).
Using techniques developed by Robert Milne of Oxford [5] one can show that there exist predicates of types:

\[ \langle \mathbb{E} \times \mathbb{E} \rangle \]
\[ \langle \mathbb{D} \times \mathbb{D} \rangle \quad \text{(one for each } \mathbb{I} \text{)} \]
\[ \langle \mathbb{E} \times \mathbb{E} \rangle \]
\[ \langle \mathbb{D} \times \mathbb{D} \rangle \]

which are directed-complete (i.e. if they hold of each member of a directed set then they hold of the union) and satisfy:

\[
\begin{align*}
\rho &< \rho' \iff \forall x. \quad \rho(x) - \rho'(x) \\
\nu &< \nu' \iff \forall \rho, \rho'. \quad [ \rho < \rho' \implies \nu(\rho[v/x]) \equiv \nu'(\rho'[v'/x]) ] \\
\rho &< \rho' \iff \forall x. \quad \rho(x) - \rho'(x) \\
\nu &< \nu' \iff \forall \rho, \rho'. \quad [ \rho < \rho' \implies \nu(\rho) \equiv \nu'(\rho) ]
\end{align*}
\]

One can then show that:

\[
\begin{align*}
\nu + \nu' &\implies \nu \circ \nu(v) \\
\nu' + \nu &\implies \nu(v) \circ \nu
\end{align*}
\]

And as it also turns out that \( \rho + \rho' \implies \rho' \circ \rho \) we have:

\[
\nu + \nu, \rho + \rho \implies \nu(\rho[v/x]) = \nu(\rho(\nu(v))/x) = \nu(\nu(v))
\]

Thus a definition of "regular" which works is given by:

**Definition 1**

\( \nu : \mathbb{E} \to \mathbb{D} \) and \( \rho : \mathbb{E} \) are regular \( \iff \nu + \nu \) and \( \rho + \rho \)

To apply this to LISP one just shows that the denotations of forms and alists are regular, this is done in section 5.

In the next section proofs of the above assertions will be given relative to the existence of the predicates. This existence (which can't be shown with the \( \text{Y} \)-operator, as the necessary functionals aren't continuous) will be proved in section 6.
3.2. Equivalent Environments

The formulation of the result about free variables also requires the use of Milne style recursive predicates viz:

\[ \Phi \subseteq \forall \rho \times \{ X | X \subseteq ld \} \]
\[ =^X \subseteq Env \times Env \quad \text{(one for each } X \subseteq ld) \]

Where intuitively \( \Phi(\nu, X) \) means the free variables of \( \nu \) are included in \( X \) and \( \rho =^X \rho' \) means \( \rho \) and \( \rho' \) "strongly" agree for all \( x \in X \). Formally we require that:

\[ \Phi(\nu, X) \iff \forall \rho, \rho', \{ X \subseteq Y \Rightarrow [ \rho =^Y \rho' \Rightarrow \nu(\rho) = \nu(\rho') ] \} \]
\[ \rho =^X \rho' \iff \forall x \in X. \rho(x) = \rho'(x) \text{ and } \Phi(\rho(x), X) \]

In section 5 below I'll show that if \( e \) is a LISP form which denotes \( G[e] \) and if \( vs(e) = \{ x | x \text{ is free in } e \} \) then \( \Phi(G[e], vs(e)) \). From this it follows (via the definition of \( \rho =^{vs(e)} \rho' \) that:

\[ \forall \rho, \rho', [ \rho =^{vs(e)} \rho' \Rightarrow G[e](\rho) = G[e](\rho') ] \]

In particular if \( e \) has no free variables then \( vs(e) = \{ \} \) and (since it's clear that for any \( \rho \) and \( \rho' ; \rho =^1 \rho' \) we have \( G[e](\rho) = G[e](\rho') \).

Somewhat less trivially: if \( \forall x \in vs(e) . \rho(x) = \rho'(x) \) and \( \rho(x) \) is a constant function (i.e. is an environment independent quantity) then again \( \rho =^{vs(e)} \rho' \) and so \( G[e](\rho) = G[e](\rho') \). This last example corresponds to the case for static binding - i.e. when objects have all their free variables bound by the time they themselves are bound. The existence of \( \Phi \) and \( =^X \) will be discussed in section 6.
4. Proofs

Readers from now on are assumed to be familiar with notations commonly employed in the literature on Mathematical Semantics.

A "domain" is a partially ordered set in which each directed subset has a least upper bound. This notion of domain is used (rather than complete lattices) for minor and nonessential technical reasons (see [1] for a discussion).

The domain intended by $Env=Id\rightarrow[Env\rightarrow D]$ is the minimal solution of the equation i.e. if $id,d$ are retracts of a universal domain (e.g. Scott's $D_y$) which represent $Id$ and $D$ respectively (in the sense that $Id=\{x|x=id(x)\}$ and $D=\{x|x=d(x)\}$) then $Y(\lambda e.id\rightarrow(e\rightarrow d))$ represents $Env$. (Here $\lambda a.b=\lambda x.\lambda y.(b(a(x)))$ and $\lambda a.b=\lambda x.\lambda y.(\text{str}(u)(a(x)))$ where $\text{str}(u)=\lambda x.(x=\bot\rightarrow x, u(x))$. From this minimality it follows that there are mappings $\lambda p.p\cdot Env\rightarrow Env$ such that:

(P1) $\bot=p_0 \in p_1 \in \ldots \in p_n \in \ldots \in p$
(P2) $p= \bot \cup p_n$
(P3) $(p_n)(m)=p_{\min\{n,m\}}$
(P4) $p_{\max}(x)p'=p(x)p'_n$

In fact if $Env$ is represented as above then $p_n=(\lambda e.id\rightarrow(e\rightarrow d))(\bot)(p)$. For $\forall \langle Env\rightarrow D \rangle

v_n$ is defined by $v_n(p)=v(p_n)$. (P4) can thus be written as: $p_{\max}(x)=p(x)_n$ and it is easy to show (see [1] for details) that: $p[v/x]_{n+1}=p_{\max}[v_n/x]$.

I shall prove $[ v\neq v' \implies v_n\neq v'(Y(F_x(v'))) ]$ by showing (by induction on $n$) that $[ v\neq v' \implies v_n\neq v'(Y(F_x(v'))) ]$ and then take a limit. Similarly $[ v\neq v' \implies Y(F_x(v))\neq v' ]$ will be proved by showing that for all $n$: $[ v\neq v' \implies F_x(v)\neq (\bot)(v)\rightarrow v' ]$. 
The following rather ad-hoc looking definition enables the clean statement of some of
the lemmas below:

**Definition 2**

\[ F: \nu_0 \to \nu_0 \text{ is "invariant at } x \text{" } \iff \forall \nu, v. \ F(v)(\rho[F(v)/x]) = v(\rho[F(v)/x]) \]

The useful applications of this definition are given in the next lemma.

**Lemma 1**

For all \( x \) \( (\lambda v. v) \) and \( (\lambda v. Y(F,v)) \) are both invariant at \( x \).

**Proof**

Trivial for \( (\lambda v. v) \), for \( (\lambda v. Y(F,v)) \) use the fixed-point property of \( Y \).

QED.

**Lemma 2**

If \( F \) is invariant at \( x \) and \( v \triangleleft v' \) then \( \forall n. v_n \triangleleft v'(v') \).

**Proof**

\( r = 0 \): Must show \( v_0 \triangleleft v'(v') \)

i.e. \( \rho \triangleleft \rho' \Rightarrow v_0(\rho[v_0/x]) \subseteq v'(\rho[F(v')/x]) \)

i.e. \( \rho \triangleleft \rho' \Rightarrow v(\bot) \subseteq v'(\rho[F(v')/x]) \)

OK as \( v \triangleleft v' \) and \( \bot \triangleleft \rho[F(v')/x] \)

\( n > 0 \): Assume true for \( n-1 \). Let \( \rho \triangleleft \rho' \). Must show \( v_n(\rho[v_n/x]) \subseteq v'(\rho[F(v')/x]) \)

i.e. \( v_n(\rho[v_{n-1}/x]) \subseteq v'(\rho[F(v')/x]) \)

need \( \rho_n[v_{n-1}/x] \triangleleft v[F(v')/x] \)

need \( v_{n-1} \triangleleft v'(v') \) - OK by induction.

QED.
Lemma 3

If $F$ is invariant at $x$ and $v = v'$ then $v \circ F(v')$

Proof

Trivial from lemma 2 as $v = \bigcup_n v_n$ and $\circ$ is directed-complete.

QED.

Lemma 4

$\forall x. \ [ v \circ v' \Rightarrow v \circ v' ]$

$\forall x. \ [ v \circ v' \Rightarrow v \circ Y(F_x(v')) ]$

Proof

Trivial consequence of lemmas 1 and 3.

QED.

Lemma 5

If $F$ is invariant at $x$ and $v = v'$ then $\forall n. F_x(v)^n(\bot) \circ F(v')$.

Proof

$n=0$: Trivial

$n>0$: Assume true for $n-1$. Need $\rho \circ \rho' \Rightarrow F_x(v)^n(\bot)(\rho[F_x(v)^n(\bot)/x]) = F(v')(\rho[F(v')/x])$

i.e. $\rho \circ \rho' \Rightarrow \nu(\rho[F_x(v)^n(\bot)/x]) = \nu(\rho[F(v')/x])$

OK if $F_x(v)^n(\bot) \circ F(v')$ - true by induction

QED.

Lemma 6

If $F$ is invariant at $x$ and $v = v'$ then $Y(F_x(v)) \circ F(v')$.

Proof

Trivial from lemma 5 as $Y(F_x(v)) = \bigcup_n F_x(v)^n(\bot)$ and $\circ$ is directed-complete.

QED.
Lemma 7

∀x. [ v∀v' =⇒ Y(F_x(v))∀v' ]
∀x. [ v∀v' =⇒ Y(F_x(v))∀Y(F_x(v')) ]

Proof

Trivial application of lemma 1 and lemma 6.
QED.

Theorem 1

If v and p are regular then v(p\[v/x\])=Y(F_x(v))p

Proof

By lemma 5 and lemma 7 we have:

\[ Y(F_x(v))∀v' \]

\[ v∀Y(F_x(v)) \]

hence from the definition of \( o^v \)

\[ Y(F_x(v))(p[Y(F_x(v))/x]) ∈ v[p[v/x]] \]

\[ v[p[v/x]] ∈ Y(F_x(v))(p[Y(F_x(v))/x]) \]

hence

\[ Y(F_x(v))(p[Y(F_x(v))/x])=v[p[v/x]] \]

Finally, using the fixed-point property of Y on the left hand side of this, we get:

\[ Y(F_x(v))p=v[p[v/x]] \]

QED.

5. Application to LISP

In this section D will be specialized to a domain appropriate for pure LISP and then the abstract results described above will be shown to hold of the denotations of LISP programs.

The semantics of LISP used here will only be described in barest outline. For further details, motivation and justification see [1] and [2].
5.1. Syntax

The syntax of LISP (as described in the manual [4] and in the notation of [9]) is given by the equations:

\[ e ::= A \mid x \mid \text{fn}[e_1;\ldots;e_n] \mid [e_1 \rightarrow e_2;\ldots;e_n \rightarrow e_{n+2}] \]

\[ \text{fn} ::= F \mid f \mid \lambda[[x_1;\ldots;x_n];e] \mid \text{label}[f;\text{fn}] \]

\[ F ::= \text{car} \mid \text{cdr} \mid \text{cons} \mid \text{atom} \mid \text{eq} \]

where the ranges of the variables \( e, A, x, f, z \) are as follows:

- \( A \) ranges over \(<\text{S-expression}>\)
- \( x, f, z \) range over \(<\text{identifier}>\)
- \( e \) ranges over \(<\text{form}>\)
- \( \text{fn} \) ranges over \(<\text{function}>\)
- \( F \) ranges over \(<\text{standard function}>\)

I use meta-variables \( x, f, z \) to range over \(<\text{identifier}>\): \( x \) is used in contexts where the identifier is a form, \( f \) where it's a function and \( z \) where it could be either.

5.2. Some Notation

In the semantics below:

\[ \text{flat}(S) = S \cup \{ \bot \} \text{ ordered by } \forall s < S. \bot \leq s. \]

\[ \lambda s_1,\ldots,s_n. E(s_1,\ldots,s_n) = \lambda s_1,\ldots,s_n. (s_1 = \bot \text{ or } s_2 = \bot \text{ or } \ldots \text{ or } s_n = \bot \rightarrow \bot, E(s_1,\ldots,s_n)) \]

\text{car, cdr, cons, atom, eq} \text{ are the appropriate functions on } S = \text{flat}(<s-S-expression>). \]

Whenever an expression \( v \) of type \( S, [\text{Env} \rightarrow S] \) or \( [\text{Env} \rightarrow \text{Funval}] \) occurs in a context requiring something of type \( [\text{Env} \rightarrow D] \) then \( v \) means (i.e should be "coerced" into) \( (\lambda p. \text{vinD}) \), \( (\lambda p. \text{v(p)} \text{ inD}) \) and \( (\lambda p. \text{v(p)} \text{ inD}) \) respectively.
5.3. Semantics

5.3.1. Denotation Domains

\[ D = S \cdot \text{Funval} \]
\[ S = \text{flat(\langle S-expression \rangle)} \]
\[ \text{Funval} = /S^* \rightarrow S/ \]

5.3.2. Environment Domain

\[ E \text{nv} = \text{Id} \rightarrow /E\text{nv} \rightarrow D/ \]

5.3.3. Semantic Functions

\[ \mathcal{G}: \text{Form} \rightarrow /E\text{nv} \rightarrow S/ \]
\[ \mathcal{G}: \text{Function} \rightarrow /E\text{nv} \rightarrow \text{Funval}/ \]

5.3.4. Semantic Equations

\[(S1) \quad \mathcal{G}[A] \rho = A \]
\[(S2) \quad \mathcal{G}[x] \rho = \rho(x) \mid S \]
\[(S3) \quad \mathcal{G}[\text{fn}[e_1;\ldots;e_n]] \rho = \mathcal{G}[\text{fn}] \rho (\mathcal{G}[e_1] \rho, \ldots, \mathcal{G}[e_n] \rho) \]
\[(S4) \quad \mathcal{G}[e_1 \rightarrow e_{12};\ldots;e_{n_1} \rightarrow e_{n_2}] \rho = (\mathcal{G}[e_1] \rho \rightarrow \mathcal{G}[e_{12}] \rho, \ldots, \mathcal{G}[e_{n_1}] \rho \rightarrow \mathcal{G}[e_{n_2}] \rho) \]
\[(S5) \quad \mathcal{G}[\text{car}] \rho = \text{car} \]
\[\mathcal{G}[\text{cdr}] \rho = \text{cdr} \]
\[\mathcal{G}[\text{cons}] \rho = \text{cons} \]
\[\mathcal{G}[\text{atom}] \rho = \text{atom} \]
\[\mathcal{G}[\text{eq}] \rho = \text{eq} \]
\[(S6) \quad \mathcal{G}[f] \rho = \rho(f) \rho | \text{Funval} \]
\[(S7) \quad \mathcal{G}[\lambda([x_1;\ldots;x_n];e)] \rho = \lambda e_1,\ldots,e_n:S.\mathcal{G}[e] \rho[e,1/\ldots][e_n/x_n] \]
\[(S8) \quad \mathcal{G}[\text{label}[f;\text{fn}]] \rho = \text{Y}(F,(\mathcal{G}[\text{fn}])) \rho \]
Theorem 2 below shows that the denotations of LISP forms and functions are regular and so Theorem 1 can be applied to them.

**Theorem 2**

\[
G[e] \times G[e] \quad \text{and} \quad \mathcal{F}[\text{fn}] \times \mathcal{F}[\text{fn}]
\]

**Proof**

A straightforward induction works. The details are as follows:
Assume \( \rho < \rho' \). I must show \( G[e] \rho \in G[e] \rho' \) and \( \mathcal{F}[\text{fn}] \rho \in \mathcal{F}[\text{fn}] \rho' \).

(1): \( G[A] \rho = \rho = A = G[e] \rho' \)

(2): \( G[x] \rho = \rho(x) \rho | S \\
G[x] \rho' = \rho'(x) \rho' | S \\
Now \( \rho < \rho' \Rightarrow \rho(x) < \rho'(x) \Rightarrow \rho(x)(\rho(x)/x) \in \rho'((\rho'(x)/x)) \Rightarrow \rho(x)(\rho(x)') \in \rho'(x)(\rho'(x)/x) \Rightarrow \rho < \rho' \) by lemma 8 below

(3): \( G[\text{fn}[e_1;..;e_n]] \rho = \mathcal{F}[\text{fn}] \rho \mathcal{F}[\text{fn}] \rho(G[e_1],..,G[e_n]) \rho \\
= \mathcal{F}[\text{fn}] \rho'(G[e_1],..,G[e_n]) \rho' \\
= G[\text{fn}[e_1;..;e_n]] \rho' \)

(4): \( G[e_1 \rightarrow e_{12};..;e_{n1} \rightarrow e_{n2}] \rho = (G[e_1] \rho \rightarrow G[e_{12}] \rho,..,G[e_{n1}] \rho \rightarrow G[e_{n2}] \rho) \\
= (G[e_1] \rho',..,G[e_{n1}] \rho' \rightarrow G[e_{n2}] \rho') \\
= G[e_1 \rightarrow e_{12};..;e_{n1} \rightarrow e_{n2}] \rho' \)

(5): \( \mathcal{F}[F] \rho = F = F = \mathcal{F}[F] \rho' \)

(6): \( \mathcal{F}[f] \rho = \rho(f) \rho | F \text{un} \\
\mathcal{F}[f] \rho' = \rho'(f) \rho' | F \text{un} \\
and \( \rho(f) \rho \in \rho'(f) \rho' \) as in (2) above.

(7): \( \mathcal{F}[\lambda[x_1;..;x_n] e] \rho = \lambda s_1;..;s_n G[e] \rho[s_1/x_1]..[s_n/x_n] \\
\mathcal{F}[\lambda[x_1;..;x_n] e] \rho' = \lambda s_1;..;s_n G[e] \rho'[s_1/x_1]..[s_n/x_n] \\
so it suffices to show \( \rho(s_1/x_1) .. [s_n/x_n] \rho'([s_1/x_1] .. [s_n/x_n]) \)

and for this it suffices to show \( \lambda \rho.(s, in D) \Rightarrow \lambda \rho.(s, in D) \)

i.e. \( \rho < \rho' \Rightarrow (\lambda \rho. e)(\rho((\lambda \rho. e)/x)) \in (\lambda \rho. e)(\rho'(\lambda \rho. e)/x) \)

i.e. \( \rho < \rho' \Rightarrow s \subseteq s' \) - which is true.
(8): \(\exists [\text{label}[f; fn]] \rho = \text{Y}(F, (\exists [fn])) \rho\)
\(\exists [\text{label}[f; fn]] \rho^* = \text{Y}(F, (\exists [fn])) \rho^*\)
hence result by lemma 7.

QED.

Lemma 8

\(\forall \rho, x. \ \rho = \rho[\rho(x)/x]\)

Proof

Follows trivially from definition of "\(\rho[\rho(x)/x]\)" and strictness of \(\rho\).

QED.

Theorem 3 below shows that if \(v_\subseteq(e)\) is the set of free variables in \(e\) then in the abstract sense discussed above the free variables of \(\Phi[e]\) "are included in" \(v_\subseteq(e)\).

The following lemma is needed for the proof. The definitions of \(\Phi\) and \(\stackrel{x}{=}\) are on page 7.

Lemma 9

(1) \(\forall v, x, Y. \ \left[ \Phi(v, X), X \subseteq Y \Rightarrow \Phi(v, Y) \right]\)

(2) \(\forall d. \ \Phi((\lambda x.d), \emptyset)\)

(3) \(\forall v, x, X. \ \left[ \Phi(v, X) \Rightarrow \Phi(\Phi(F, \lambda (v)), X \setminus \{x\}) \right]\)

Proof

(1): Trivial.

(2): Trivial.

(3): I show \(\Phi(v, X) \Rightarrow \Phi(F, \lambda (v)), X \setminus \{x\}\) by induction on \(n\). Assume \(\Phi(v, X)\).

\(n=0: \Phi(\bot, X \setminus \{x\})\) is clearly true.
n>0: Assume true for n-1.
\[ \Phi(F_n(v)(\bot),X\{x\}) \iff p=^x(x)p' \Rightarrow F_n(v)(\bot)p=F_n(v)(\bot)p' \]
\[ \iff p=^x(x)p' \Rightarrow v(p[F_n(v)(\bot)/x])=v(p'[F_n(v)(\bot)/x]). \]
\[ \iff p=^x(x)p' \Rightarrow F_n(v)(\bot)x=xp'=\Phi(F_n(v)(\bot)/x). \]
which is true by induction and (1) above.

QED.

Theorem 3
\[ \forall e<\text{form}. \quad \Phi(\gamma[e]_v(e)) \]
\[ \forall e<\text{function}. \quad \Phi(\gamma[f]_v(e),v(e)) \]

Proof
A straightforward structural induction works. Let vs(e)\subseteq X.

\( e=x: \)
Must show \( p=^x(x)p' \Rightarrow p(x)p=p'*(x)p' \). Now vs(e)=\{x\} \subseteq X so if \( p=^x(x)p' \):
\( p(x)=p'*(x) \) and \( \Phi(p(x),X) \) hence \( p(x)p=p(x)p'=p'(x)p' \).

\( e=A: \)
Must show \( p=^x(x)p' \Rightarrow (\Phi(A)p=\gamma[A]_v(e)) \) - which is clearly true.

\( e=fn[e_1;\ldots;e_n]: \)
we have by induction that \( \Phi(\gamma[fn]_v(e),v(e)) \) and \( \Phi(\gamma[e]_v(e),v(e)) \).
Hence by lemma 9 \( \Phi(\gamma[fn]_v(e),X) \) and \( \Phi(\gamma[e]_v(e),X) \) as \( vs(fn),v(e)\subseteq v(e)\subseteq X \).
So if \( p=^x(x)p' \) then \( \gamma[fn]_v(e)p=\gamma[fn]_v(e)p' \) and \( \gamma[e]_v(e)p=\gamma[e]_v(e)p' \)
and hence \( \Phi(\gamma[e]_v(e),p=\gamma[A]_v(e)p'. \)

\( e=[e_1\rightarrow e_2;\ldots;e_n\rightarrow e_n]: \)
Argument as above.

Now let \( vs(fn)\subseteq X \).

\( fn=f: \)
Similar to "e=x" case above.

\( fn=F: \)
Similar to "e=A" case above.
fn=\lambda ([x_1;...;x_n]; e):
\Phi_5([\lambda ([x_1;...;x_n]; e)]\rho = \lambda e_1;...;e_n.e_5.e)[e]\rho[e_1/x_1]...[e_n/x_n]
vs(fn)=vs(e)\{x_1;...;x_n\} so vs(e)\subset XU\{x_1;...;x_n\}.
Now by lemma 9(1,2) if Y=XU\{x_1;...;x_n\} then
p=\pi \Rightarrow p[s_1/x_1]...[s_n/x_n] \Rightarrow p'[s_1/x_1]...[s_n/x_n]
so as \Phi(\Phi([e]),vs(e)); \Phi([e],p[s_1/x_1]...[s_n/x_n],p'[s_1/x_1]...[s_n/x_n] = \Phi([e],p'=p)[s_1/x_1]...[s_n/x_n].

fn=label[f;fn_1];
We have by induction \Phi(\Phi([fn_1]),[vs(fn_1)]) where vs(fn_1)\{f\}=vs(fn)\subset X.
So by lemma 9(3) and induction \Phi(\Phi([fn]),[vs(fn_1)\{f\})
hence \Phi(\Phi([fn]),X).
Q.E.D.

As an application one can show that adding new definitions to an environment doesn't change the values of the old ones as long as previously used variables aren't overwritten. This is an important lemma needed in proving the correctness of eval.
Here it's a trivial consequence of Theorem 3 but originally (see [1]) it needed a long ad-hoc proof which confused general arguments with LISP specific ones. To see how it follows consider an environment \rho which defines a set of functions all of whose free variables are included in X\subset I\rho. Suppose x is a new function not included in X. We wish to show that if e is a form (or function) then as long as vs(e)\subset X (i.e. e only uses the old functions) we have for any v: \Phi([e],\rho=\Phi([e],\rho[v/x]). But this is now trivial for \Phi([e],X) and \rho=\pi \Rightarrow \rho[v/x]. Saying this formally yields the following theorem (in which "\rho[v/x]" above is replaced by "\rho[v/x]").
Theorem 4

Suppose $\rho, \rho' \langle Env, e \langle form \rangle$ are such that for some $X \subseteq I d$ we have:

1. $\forall x \in X. \exists fn, \langle function \rangle. \rho(x) = \rho'(x) = \Theta[fn, t] \land vs(fn) \subseteq X.$
2. $vs(e) \subseteq X$

then $\Gamma[e] \rho = \Gamma[e] \rho'.$

Proof

By theorem 3 $\Phi(\Gamma[e], X)$ and $\rho = \gamma \rho'.$ The result follows from the definition of $\Phi.$

QED.

6. Existence of Predicates

In all the above the existence of the predicates $\alpha, \alpha^*, \gamma, \gamma^* = X$ has been assumed. However this existence cannot be deduced immediately from the recursive definitions as the predicates being defined aren't necessarily monotonic. The existence proofs to be described are directly based on techniques developed by Robert Milne [5]. Similar methods have recently been independently discovered by Reynolds [7]. For the current purposes it's only necessary to know that the required predicates exist, however Milne's work shows one can expect them to be unique also. I haven't checked this for the predicates used here.

We define by induction on $n$ predicates:

$\alpha_n \subseteq Env \times Env$

$\alpha^*_n \subseteq V_D \times V_D$

and then set:

\[ \rho \circ \rho' \iff \forall n. \rho \circ \rho' \]

\[ \nu \circ\nu' \iff \forall n. \nu \circ\nu' \]

it follows (details below) that \( \sigma, \sigma' \) satisfy the desired equations and are directed-complete.

**Definition 3**

\[ \rho \circ \rho' \iff \forall x. \rho(x) \circ \rho'(x) \]

\[ \nu \circ\nu' \iff \forall \left( [v_0/x] \right) \subseteq \nu'(\left[ v'/x \right]) \]

\[ \nu \circ\nu_{v,v'} \iff \forall \rho, \rho'. \left[ \rho \circ \rho' \Rightarrow \nu, \nu' \right] \quad \left[ \rho \circ \rho' \Rightarrow \nu, \nu' \right] \]

The following two lemmas are needed to prove Theorem 5 below.

**Lemma 10**

(1) \( \forall \rho, \rho'. \left[ \rho \circ \rho' \Rightarrow \rho \circ \rho' \right] \)

(2) \( \forall \rho, \rho'. \left[ \rho \circ \rho' \Rightarrow \rho \circ \rho' \right] \)

(3) \( \forall v, v'. \left[ \nu \circ \nu' \Rightarrow \nu \circ \nu' \right] \)

(4) \( \forall v, v'. \left[ \nu \circ \nu' \Rightarrow \nu \circ \nu' \right] \)

**Proof**

I show that (3) \( \Rightarrow (1) \), (4) \( \Rightarrow (2) \), (2) \( \Rightarrow (3) \), (1) \( \Rightarrow (4) \)

(3) \( \text{Must show } \nu \circ \nu' \Rightarrow \nu \circ \nu' \). Clearly \( \perp \circ \perp \) and we have:

\[ \nu \circ \nu', \perp \circ \perp \Rightarrow \nu, \left( \perp [v/x] \right) \subseteq \nu'(\left[ v'/x \right]) \]

\[ \Rightarrow \nu, \left( \perp [v_0/x] \right) \subseteq \nu'(\left[ v'/x \right]) \]

\[ \Leftarrow \nu \circ \nu' \]
(4_0): Must show \( v \circ_0 v^* \Rightarrow v \circ_1^* v^* \).
Assume \( v \circ_0 v^* \) and \( \rho \circ_0 \rho^* \).
Must show \( v_0 ((\rho_0[x]) \equiv v^*(\rho^*[v^*/x])) \)
i.e. \( v(\bot) \equiv v^*(\rho^*[v^*/x]) \)
but \( v(\bot) \equiv v((\bot)[v_0/x]) \equiv v^*(\bot[v^*/x]) \equiv v^*(\rho^*[v^*/x]) \).

\((3_n) = (1_n): \) Assume \((3_n)\). To show \((1_n)\) let \( \rho \circ_{n,1} \rho^* \).
Must show \( \rho \circ_{n,1} \rho^* \) i.e. \( \forall x. \rho(x) \circ_{n,1} \rho^*(x) \).
But if \( \rho \circ_{n,1} \rho^* \) then \( \forall x. \rho(x) \circ_{n,1} \rho^*(x) \) so \( \forall x. \rho(x) \circ_{n,1} \rho^*(x) \) by \((3_n)\).

\((\ell_n) = (2_n): \) \( \rho \circ_{n,1} \rho^* \Leftrightarrow \forall x. \rho(x) \circ_{n,1} \rho^*(x) \)
\( \Rightarrow \forall x. \rho(x) \circ_{n,1} \rho^*(x) \) by \((4_n)\)
\( \Rightarrow \forall x. \rho_{n,1}(x) \circ_{n,1} \rho^*(x) \)
\( \Rightarrow \forall x. \rho_{n,1} \circ_{n,1} \rho^* \)

\((2_n,1) = (3_n): \) Assume \((2_n,1)\). To show \((3_n)\) let \( v \circ_{n,1} v^* \) and \( \rho \circ_{n,1} \rho^* \).
Then \( \rho \circ_{n,1} \rho^* \) from \((2_n,1)\).
So \( v_{n,1}(\rho_{n,1}[v/x]) \equiv v^*(\rho^*[v^*/x]) \)
i.e. \( v(\rho_{n,1}[v/x]) \equiv v^*(\rho^*[v^*/x]) \)
hence \( v_n(\rho[v/x]) = v(\rho_{n,1}[v_{n,1}/x]) \)
\( \equiv v(\rho_{n,1}[v_{n,1}/x]) \)
\( \equiv v^*(\rho^*[v^*/x]) \).

\((\ell_n,1) = (4_n): \) Assume \((1_n,1)\). To show \((4_n)\) let \( v \circ_{n,1} v^* \) and \( \rho \circ_n \rho^* \).
Then \( \rho \circ_n \rho^* \) so \( v_n(\rho_{n,1}[v/x]) \equiv v^*(\rho^*[v^*/x]) \)
hence \( (v_n)_{n,1}(\rho_{n,1}[v_{n,1}/x]) = v(\rho_{n,1}[v_{n,1}/x]) \)
\( = v_n(\rho[v/x]) \)
\( \equiv v^*(\rho^*[v^*/x]) \).

QED.
Lemma 11

If \{v_\alpha\} is directed then \[[\forall_\alpha. v_\alpha \alpha^*_{\alpha}, v^*]\] \Rightarrow \[(\bigcup_\alpha v_\alpha) \alpha^*_{\alpha}, v^*\] .

Proof

Cases on n:

n=0: \nu_\alpha \nu_\alpha^*_{\alpha} \Leftrightarrow v_\alpha(\nu[\nu_\alpha /x]) \in v^*(\nu'[\nu' /x])
so \bigcup_\alpha v_\alpha(\nu[\bigcup_\alpha v_\alpha /x]) \in v^*(\nu'[\nu' /x]).

n>0: Let \nu_\alpha^*_{\alpha}, \nu' then \forall_\alpha. \forall_\alpha(\nu[\nu_\alpha /x]) \in v^*(\nu'[\nu' /x])
so \bigcup_\alpha v_\alpha(\nu[\bigcup_\alpha v_\alpha /x]) \in v^*(\nu'[\nu' /x])

hence \bigcup_\alpha v_\alpha \nu^*_{\alpha}, v^*.

Q.E.D.

Theorem 5

\alpha and \alpha^* are directed-complete and satisfy:

\nu_\alpha \nu_\alpha^*_{\alpha} \nu^* \Leftrightarrow \forall_\alpha. \nu_\alpha(\alpha^*_{\alpha} \nu^*) \equiv v^*(\alpha^*_{\alpha} \nu^*)
\nu_\alpha \omega^* \nu^* \Leftrightarrow \forall_\alpha. \nu_\alpha \nu^*(\nu \equiv v^*(\nu)^{\nu^* /x})

Proof

To show \alpha^* directed-complete we have:

\forall_\alpha. \forall_\alpha \nu_\alpha \omega^* \nu^* \Leftrightarrow \forall_\alpha \forall_\alpha \nu_\alpha \omega^* \nu^*
\Leftrightarrow \forall_\alpha. \forall_\alpha \nu_\alpha \omega^* \nu^*
\Leftrightarrow \forall_\alpha. \bigcup_\alpha v_\alpha \omega^* \nu^* \text{ by lemma } 11
\Leftrightarrow \bigcup_\alpha v_\alpha \omega^* \nu^*

Showing \[(\forall_\alpha. \nu_\alpha \omega^* \nu^*) \Rightarrow [\nu_\alpha(\bigcup_\alpha v_\alpha)]\] is trivial.

The directed-completeness of \alpha follows directly from its definition and the directed-completeness of \alpha^* for all \alpha.
To prove the rest of the theorem we have:

\[ \rho \circ \rho' \iff \forall x. \rho(x) \circ \rho'(x) \]

To show \( v \circ v' \Rightarrow \forall \rho, \rho'. [\rho \circ \rho' \Rightarrow v(\rho[v/x]) \in v'(\rho'[v'/x])] \) assume \( v \circ v' \) and \( \rho \circ \rho' \). Then \( v \circ v', \rho \circ \rho' \)
so \( v_n(\rho[v/x]) \in v'(\rho'[v'/x]) \)
hence unioning over \( n \): \( v(\rho[v/x]) \in v'(\rho'[v'/x]) \).

To show \( \forall \rho, \rho'. [\rho \circ \rho' \Rightarrow \rho \circ \rho] \Rightarrow v(\rho[v/x]) \in v'(\rho'[v'/x]) \) assume \( \rho \circ \rho' \Rightarrow v(\rho[v/x]) \in v'(\rho'[v'/x]) \).
show \( \forall n. v \circ v, \rho, \rho' \) by induction on \( n \).

\( n=0 \): \( \bot \circ \bot \Rightarrow \bot \) so \( \bot \in \bot \Rightarrow \bot \Rightarrow \bot \in \bot \Rightarrow \bot \in \bot \Rightarrow \bot \Rightarrow \bot \)
so \( v \circ v \).

\( n>0 \): By lemma 10: \( \rho \circ \rho' \Rightarrow \rho \circ \rho \Rightarrow \rho \circ \rho \Rightarrow \forall m. \rho \circ \rho \).
so \( \rho \circ \rho' \).
Hence \( v(\rho, \rho' \Rightarrow v(\rho'[v'/x]) \) so \( v_n(\rho[v/x]) \in v(\rho'[v'/x]) \).
Thus \( v \circ v \).

So \( \forall n. v \circ v, \rho, \rho' \) and hence \( v \circ v \).

QED.

The construction of \( \Phi \) and \( =^X \) is very similar to the construction above. As before we start by defining "finite" approximations to the relations viz.

**Definition 4**

\[ \Phi_n(v, X) \iff \forall Y, \rho, \rho'. [X \subseteq Y \Rightarrow [\rho = X, \rho' \Rightarrow v_n(\rho) = v_n(\rho')] ] \]

\[ \rho =^X_0 \rho' = true \]
\[ \rho =^X_{n+1} \rho' \iff \forall x(X, \rho(x) = \rho'(x) \text{ and } \Phi_n(\rho(x), X)) \]

We then prove a lemma similar to lemma 10 viz.
Lemma 12

(1) $\forall \nu, X. \left[ \Phi_{n-1}(\nu, X) \Rightarrow \Phi_n(\nu, X) \right]$

(2) $\forall \nu, X. \left[ \Phi_n(\nu, X) \Rightarrow \Phi_{n-1}(\nu, X) \right]$

(3) $\forall \rho, \rho', X. \left[ \rho =_{\nu, n} \rho' \Rightarrow \rho =_{\nu, n} \rho' \right]$

(4) $\forall \rho, \rho', X. \left[ \rho =_{\nu, n} \rho' \Rightarrow \rho =_{\nu, n} \rho' \right]$

Proof

Same as lemma 10 (mutatis mutandis).

QED.

From this it follows that if we define $\Phi$ and $=^X$ by:

$\Phi(\nu, X) \iff \forall n. \Phi_n(\nu, X)$

$\rho =^X \rho' \iff \forall n. \rho =_{\nu, n} \rho'$

then $\Phi$ and $=^X$ have the desired properties.

7. Concluding Remarks

We have presented above a partial axiomatization of dynamic binding. What has been shown is that if $\nu \in \mathcal{E}_{nu \rightarrow D}$ satisfies $\nu \cdot \nu$ (i.e. is regular) and $\Phi(\nu, X)$ for some $X \subseteq I_d$ then useful theorems follow. What is left open is just how many other axioms will eventually be required. To answer this we need first to know which theorems we want and to answer this we must attack "real" problems such as the correctness of compilers and interpreters. Doing this should reveal the general theorems about dynamic binding that must follow from any adequate theory.
The theorems proved here are not yet general enough. For example if we consider
the obvious extension of the semantics to handle funargs (see [1]) then the proofs
that \( \text{G}[e] \) and \( \text{G}[\text{fn}] \) are regular fail. In fact by replacing
the occurrences of "\( \equiv \)" in the definitions of \( <, >^* \) and + by another predicate
(which needs to be defined recursively) it's easy to cover this case. Unfortunately I
don't see a uniform way of defining \( <, >^* \) and + to cover all useful \( D \).

Having to separately prove the existence of all predicates is a big nuisance. One
step toward a general justification of recursive predicate definitions has been
provided by Milne and Reynolds. Both give uniform accounts of how to construct
recursive predicates from their defining equations. In fact the constructions given
above are (more or less) instances of Milne's techniques. It would help a lot if
syntactic criteria on definitions could be developed to decide if the things purported
to be defined actually exist. Milne [private communication] has made progress toward
this by analysing the structure of some of the expressions which occur in definitions
and showing that these legitimate instances of his general construction.

It's clear that many of the above proofs can't be done in existing formalisms (e.g. LCF)
- the required predicates can't be defined in them. One way to fix this would be
to develop extensions, another would be to develop a translator from proofs using
predicates to proofs which don't. The latter probably won't be adequate because
theorems may require the use of predicates in their statement at the general level
(even if all their useful instances don't).
8. References


