Functional Programming

Evaluation and Typing

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Results of computations

Normal forms

- expensive to compute
- rarely needed for evaluation
- key to expense: evaluation under lambda

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Definition: Weak head-normal form

A pure lambda term is in **weak head-normal form** (or a **value**) iff it has the form

$$V ::= \lambda x.M$$

All other terms are **non-values**.

Deterministic Evaluation

Definition

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Reduction strategy: Call-by-value (based on β-value reduction)

$$(\lambda x.M)\ V \rightarrow_{\beta V} M[x \mapsto V] \qquad \frac{M \rightarrow_{\beta V} M'}{(M\ N) \rightarrow_{\beta V} (M'\ N)} \qquad \frac{N \rightarrow_{\beta V} N'}{(V\ N) \rightarrow_{\beta V} (V\ N')}$$

Evaluation in Functional Programming Languages

Evaluation of closed terms

- Reduction is restricted to **closed terms** and stops at weak head-normal forms.

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- Consequence: substitution need not rename variables!
- Substitution can be avoided by implementation tricks
- Datatypes are not encoded, but added to the calculus

Applied Lambda Calculus

Syntax

Add constants as values to the pure lambda calculus

$$\begin{aligned} L, M, N &::= x \mid \lambda x.M \mid M\,N \mid C \\ C &::= \mathit{TRUE} \mid \mathit{FALSE} \mid \mathit{IF} \mid 0 \mid 1 \mid \cdots \mid \mathit{SUCC} \mid \cdots \mid \mathit{PAIR} \mid \mathit{FST} \mid \mathit{SND} \\ V, W &::= \lambda x.M \mid C \mid \mathit{IF}\,V \mid \mathit{IF}\,V\,M \mid \mathit{PAIR}\,M \mid \mathit{PAIR}\,M\,N \end{aligned}$$

Semantics (call-by-name)

β reduction and δ **reduction rules** for the constants

$$(\lambda x.M)\,N \rightarrow_\beta M[x \mapsto N] \qquad \mathit{IF}\,\mathit{TRUE}\,M\,N \rightarrow_\delta M \qquad \mathit{IF}\,\mathit{FALSE}\,M\,N \rightarrow_\delta N$$

$$\mathit{FST}\,(\mathit{PAIR}\,M\,N) \rightarrow_\delta M \qquad \mathit{SND}\,(\mathit{PAIR}\,M\,N) \rightarrow_\delta N$$

Handling Constants

Context rule for constants

$$\frac{N \rightarrow_x N' \qquad V \in \{IF, FST, SND, SUCC\}}{(V\ \ N) \rightarrow_x (V\ \ N')}$$

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New source of errors: stuck terms

- Pure lambda calculus: closed term is either value or can reduce
- Applied lambda calculus: there are closed non-value terms that cannot be reduced
  - $TRUE\,V$
  - $IF\,(\lambda x.L)\,M\,N$
  - $FST\,(\lambda x.M)$
  - $IF\,(PAIR\,M\,N)\,M'\,N'$
  - ...
- These terms are **stuck terms**

Typing

Typing rules out stuck terms

What is a typing?

- Typing $M : T$ is a relation between terms and types
- Typing characterizes terms with a certain behavior

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Simple types for the applied lambda calculus

$$T ::= T \to T \mid \mathit{Nat} \mid \mathit{Bool} \mid \mathit{Pair}\, T\ T$$

Intended Behavior

If $M : T$, then M is not stuck.

Simple Types

Definition: Typing assumption (environment)

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Definition: Typing rules — lambda calculus with numbers

$$\textsc{Var}\quad A, x : T, A' \vdash x : T$$

$$\textsc{Lam}\quad \frac{A, x : T \vdash M : T'}{A \vdash \lambda x.M : T \rightarrow T'}$$

$$\textsc{App}\quad \frac{A \vdash M : T \rightarrow T' \qquad A \vdash N : T}{A \vdash M\,N : T'}$$

$$\textsc{Num}\quad A \vdash n : \mathit{Nat}$$

$$\textsc{Succ}\quad \frac{A \vdash M : \mathit{Nat}}{A \vdash \mathit{SUCC}\,M : \mathit{Nat}}$$

More typing rules

Definition: Typing rules — boolean fragment

$$\textsc{True}\quad \frac{}{A \vdash TRUE : Bool}$$

$$\textsc{False}\quad \frac{}{A \vdash FALSE : Bool}$$

$$\textsc{If}\quad \frac{A \vdash L : Bool \qquad A \vdash M : T \qquad A \vdash N : T}{A \vdash IF\, L\, M\, N : T}$$

More typing rules

Definition: Typing rules — boolean fragment

$$\textsc{True}\quad \overline{A \vdash \mathit{TRUE} : \mathit{Bool}}$$

$$\textsc{False}\quad \overline{A \vdash \mathit{FALSE} : \mathit{Bool}}$$

$$\textsc{If}\quad \frac{A \vdash L : \mathit{Bool} \qquad A \vdash M : T \qquad A \vdash N : T}{A \vdash \mathit{IF}\, L\, M\, N : T}$$

Definition: Typing rules — pairs

$$\textsc{Pair}\quad \frac{A \vdash M : T \qquad A \vdash N : T'}{A \vdash \mathit{PAIR}\, M\, N : \mathit{Pair}\, T\, T'}$$

$$\textsc{Fst}\quad \frac{A \vdash M : \mathit{Pair}\, T\, T'}{A \vdash \mathit{FST}\, M : T}$$

$$\textsc{Snd}\quad \frac{A \vdash M : \mathit{Pair}\, T\, T'}{A \vdash \mathit{SND}\, M : T'}$$

Example Inference Tree

$$
\dfrac{\dfrac{\dfrac{\cdots \vdash f : \alpha \to \alpha \qquad \dfrac{\cdots \vdash f : \alpha \to \alpha \qquad \cdots \vdash x : \alpha}{\cdots \vdash f\, x : \alpha}}{f : \alpha \to \alpha, x : \alpha \vdash f\,(f\, x) : \alpha}}{f : \alpha \to \alpha \vdash \lambda x.f\,(f\, x) : \alpha \to \alpha}}{\cdot \vdash \lambda f.\lambda x.f\,(f\, x) : (\alpha \to \alpha) \to \alpha \to \alpha}
$$

Type Soundness

Type Preservation

If $\cdot \vdash M : T$ and $M \rightarrow N$, then $\cdot \vdash N : T$.

Proof by induction on $M \rightarrow N$.

Progress

If $\cdot \vdash M : T$, then either M is a value or there exists M' such that $M \rightarrow M'$.

Proof by induction on $A \vdash M : T$.

Type Soundness

If $\cdot \vdash M : T$, then either

1. exists V such that $M \rightarrow^* V$ or
2. for each N, such that $M \rightarrow^* N$ there exists N' such that $N \rightarrow N'$.

Type Inference for the Simply-Typed Lambda Calculus

Type Inference for the Simply-Typed Lambda Calculus (STLC)

Typing Problems

- Type checking: Given environment A, a term M and a type T, is $A \vdash M : T$ derivable?
- Type inference: Given a term M, are there A and T such that $A \vdash M : T$ is derivable?

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Typing Problems for STLC

- Type checking and type inference are decidable for STLC
- Moreover, for each typable M there is a **principal typing** $A \vdash M : T$ such that any other typing is a substitution instance of the principal typing.

Prerequisites for Type Inference for STLC

Unification

Let $\mathcal{E}$ be a set of equations on types.

Unifiers and Most General Unifiers

- A substitution S is a **unifier of** $\mathcal{E}$ if, for each $T \doteq T' \in \mathcal{E}$, it holds that $ST = ST'$.
- A substitution S is a **most general unifier of** $\mathcal{E}$ if S is a unifier of $\mathcal{E}$ and for every other unifier S' of $\mathcal{E}$, there is a substitution U such that $S' = U \circ S$.

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Unification

There is an algorithm $\mathcal{U}$ that, on input $\mathcal{E}$, either returns a most general unifier of $\mathcal{E}$ or fails if none exists.

Principal Type Inference for STLC

The algorithm (due to John Mitchell) transforms a term into a principal typing judgment for the term or fails if no typing exists.

$$
\begin{array}{lcl}
\mathcal{P}(x) & = & \textbf{return } \lceil x : \alpha \vdash x : \alpha \rceil \\
\mathcal{P}(\lambda x.M) & = & \textbf{let } \lceil A \vdash M : T \rceil \leftarrow \mathcal{P}(M) \textbf{ in} \\
& & \textbf{if } A \textbf{ has form } A', x : T_x \textbf{ then return } \lceil A' \vdash \lambda x.M : T_x \to T \rceil \\
& & \textbf{else choose } \alpha \notin \mathrm{var}\,(A, T) \textbf{ in} \\
& & \qquad \textbf{return } \lceil A \vdash \lambda x.M : \alpha \to T \rceil \\
\mathcal{P}(M_0\, M_1) & = & \textbf{let } \lceil A_0 \vdash M_0 : T_0 \rceil \leftarrow \mathcal{P}(M_0) \textbf{ in} \\
& & \textbf{let } \lceil A_1 \vdash M_1 : T_1 \rceil \leftarrow \mathcal{P}(M_1) \textbf{ in} \\
& & \textbf{with disjoint type variables in } (A_0, T_0) \textbf{ and } (A_1, T_1) \\
& & \textbf{choose } \alpha \notin \mathrm{var}\,(A_0, A_1, T_0, T_1) \textbf{ in} \\
& & \textbf{let } S \leftarrow \mathcal{U}(A_0 \doteq A_1, T_0 \doteq T_1 \to \alpha) \textbf{ in} \\
& & \textbf{return } \lceil SA_0 \cup SA_1 \vdash M_0\, M_1 : S\alpha \rceil \\
\mathcal{P}(n) & = & \textbf{return } \lceil \cdot \vdash n : Nat \rceil \\
\mathcal{P}(SUCC\, M) & = & \textbf{let } \lceil A \vdash M : T \rceil \leftarrow \mathcal{P}(M) \textbf{ in} \\
& & \textbf{let } S \leftarrow \mathcal{U}(T \doteq Nat) \textbf{ in} \\
& & \textbf{return } \lceil SA \vdash SUCC\, M : Nat \rceil
\end{array}
$$

Properties of Type Inference

Soundness

If $\mathcal{P}(M) = \lceil A \vdash M : T \rceil$, then $A \vdash M : T$ is derivable.

Completeness

If $A \vdash M : T$ is derivable, then $\mathcal{P}(M)$ succeeds with result $\lceil A' \vdash M : T' \rceil$ such that $A = SA'$ and $T = ST'$ for some substitution S.

Wrapup

- Call-by-name and call-by-value are deterministic evaluation strategies that are more efficient than full β reduction
- Applied lambda calculus contains constants that encode operations on datatypes
- Applied lambda calculus can have stuck terms
- Simple types avoid stuck terms
- Type checking and type inference for simple types is decidable
- There is a sound and complete algorithm for type inference for simple types