

## 1 Recap

Last lecture we saw how to unify types.

$$\begin{aligned} \text{Unify}(\emptyset) &\triangleq I \\ \text{Unify}(\alpha = \alpha, E) &\triangleq \text{Unify}(E) \\ \text{Unify}(\alpha = \tau, E) &\triangleq \{\tau/\alpha\} \cdot \text{Unify}(E\{\tau/\alpha\}), \quad \alpha \notin \text{FV}(\tau) \\ \text{Unify}(\sigma_1 \rightarrow \tau_1 = \sigma_2 \rightarrow \tau_2, E) &\triangleq \text{Unify}(\sigma_1 = \sigma_2, \tau_1 = \tau_2, E) \end{aligned}$$

where  $I$  is the identity substitution  $\alpha \mapsto \alpha$ . Substitutions are applied from left to right, so the composition  $ST$  means: do  $S$  first, then do  $T$ .

## 2 Polymorphic $\lambda$ -Calculus

Suppose we have base types `int` and `bool`. The problem with the simple type inference mechanism that we have presented is that we do not have quite as much *polymorphism*<sup>1</sup> as we would like. For example, consider a program that binds a variable to the identity function, then applies it to an `int` and also to a `bool`.

$$\text{let } f = \lambda x. x \text{ in if } (f \text{ true}) \text{ then } (f \text{ 3}) \text{ else } (f \text{ 4}) \quad (1)$$

The type checker encounters the `bool` first and says that the function is of type `bool`  $\rightarrow$  `bool`, then gives an error when it sees the `int` parameter, whereas we really want it to be interpreted as type `bool`  $\rightarrow$  `bool` when applied to a `bool` parameter and `int`  $\rightarrow$  `int` when applied to an `int` parameter.

We can handle this by introducing a new type constructor that quantifies over types.

$$\tau ::= \text{int} \mid \text{bool} \mid \alpha \mid \sigma \rightarrow \tau \mid \forall \alpha. \tau \quad (2)$$

The type  $\forall \alpha. \tau$  can be viewed as a *polymorphic type* or *type schema*, a pattern with type variables that can be instantiated to obtain actual types. For example, the polymorphic type of the identity function will be the type schema

$$\forall \alpha. \alpha \rightarrow \alpha$$

and the type of the  $K$  combinator  $\lambda xy. x$  will be

$$\forall \alpha. \forall \beta. \alpha \rightarrow \beta \rightarrow \alpha.$$

There will be rules that allow us to delay the instantiation of the type variables until the function is applied. Thus we can interpret the identity function as `int`  $\rightarrow$  `int` or `bool`  $\rightarrow$  `bool` depending on context.

The resulting language is called the *polymorphic  $\lambda$ -calculus*. In this new language, the terms and evaluation rules are the same, but the types are defined by (??). All the terms that were previously well-typed will still be well-typed, but there will be more well-typed terms than before; for example, (??).

<sup>1</sup>Greek for “many forms”

### 3 Typing Rules

In addition to the old typing rules

$$\begin{array}{c} \Gamma \vdash n : \text{int} \quad (\text{and similarly for other constants}) \\ \\ \frac{\Gamma \vdash e : \sigma \rightarrow \tau \quad \Gamma \vdash d : \sigma}{\Gamma \vdash (e \ d) : \tau} \end{array} \qquad \begin{array}{c} \Gamma, x : \tau \vdash x : \tau \\ \\ \frac{\Gamma, x : \sigma \vdash e : \tau}{\Gamma \vdash \lambda x. e : \sigma \rightarrow \tau} \end{array}$$

we add the following two new rules for polymorphic types:

$$\frac{\Gamma \vdash e : \tau}{\Gamma \vdash e : \forall \alpha. \tau} \quad (\alpha \notin FV(\Gamma)) \qquad \frac{\Gamma \vdash e : \forall \alpha. \tau}{\Gamma \vdash e : \tau \{ \sigma / \alpha \}}$$

These are called the *generalization rule* and the *instantiation rule*, respectively.

The notation  $\tau \{ \sigma / \alpha \}$  refers to the safe substitution of the type  $\sigma$  for the type variable  $\alpha$  in  $\tau$ . Here the binding operator  $\forall \alpha$  binds the type variable  $\alpha$  in the same way that  $\lambda x$  binds the variable  $x$  in  $\lambda$ -terms, and the notions of scope, free and bound variables are the same. In particular, one can  $\alpha$ -convert type variables as necessary to avoid the capture of free type variables when performing substitutions.

The generalization rule includes the side condition  $\alpha \notin FV(\Gamma)$ . The idea here is that the type judgment  $\Gamma \vdash e : \tau$  must hold without any assumptions involving  $\alpha$ ; if so, then we can conclude that  $\alpha$  could have been any type  $\sigma$ , and the type judgment  $\Gamma \vdash e : \tau \{ \sigma / \alpha \}$  would also hold.

### 4 Examples

Here is a derivation of the polymorphic type of  $K$  in this system.

$$\frac{\frac{\frac{\frac{\frac{\vdash \lambda x. \lambda y. x : \forall \alpha. \forall \beta. \alpha \rightarrow \beta \rightarrow \alpha}{\vdash \lambda x. \lambda y. x : \forall \beta. \alpha \rightarrow \beta \rightarrow \alpha}}{\vdash \lambda x. \lambda y. x : \alpha \rightarrow \beta \rightarrow \alpha}}{x : \alpha \vdash \lambda y. x : \beta \rightarrow \alpha}}{x : \alpha, y : \beta \vdash x : \alpha}}{\vdash \lambda x. \lambda y. x : \forall \alpha. \forall \beta. \alpha \rightarrow \beta \rightarrow \alpha}}$$

Starting from  $x : \alpha, y : \beta \vdash x : \alpha$ , two applications of the abstraction rule yield  $\vdash \lambda x. \lambda y. x : \alpha \rightarrow \beta \rightarrow \alpha$ , then two applications of the generalization rule yield  $\vdash \lambda x. \lambda y. x : \forall \alpha. \forall \beta. \alpha \rightarrow \beta \rightarrow \alpha$ .

Some terms are typable in this system that were not typable before. For example, the term  $\lambda x. xx$  is typable:

$$\frac{\frac{\frac{\frac{\vdash \lambda x. xx : \forall \beta. (\forall \alpha. \alpha) \rightarrow \beta}{\vdash \lambda x. xx : (\forall \alpha. \alpha) \rightarrow \beta}}{x : \forall \alpha. \alpha \vdash xx : \beta}}{x : \forall \alpha. \alpha \vdash x : \alpha \rightarrow \beta} \quad \frac{x : \forall \alpha. \alpha \vdash x : \alpha}{x : \forall \alpha. \alpha \vdash x : \forall \alpha. \alpha}}{x : \forall \alpha. \alpha \vdash x : \forall \alpha. \alpha \quad x : \forall \alpha. \alpha \vdash x : \forall \alpha. \alpha}}$$

Unfortunately, this type is not too meaningful, because *nothing* has type  $\forall \alpha. \alpha$ . This type is said to be *uninhabited*, and we give it a name: **Void**. However, by a similar argument, we can show that  $\lambda x. xx$  also has type  $\forall \beta. (\forall \alpha. \alpha \rightarrow \alpha) \rightarrow (\beta \rightarrow \beta)$ , which *is* meaningful.

$$\begin{array}{c}
\vdash \lambda x. xx : \forall \beta. (\forall \alpha. \alpha \rightarrow \alpha) \rightarrow (\beta \rightarrow \beta) \\
\hline
\vdash \lambda x. xx : (\forall \alpha. \alpha \rightarrow \alpha) \rightarrow (\beta \rightarrow \beta) \\
\hline
x : \forall \alpha. \alpha \rightarrow \alpha \vdash xx : \beta \rightarrow \beta \\
\hline
x : \forall \alpha. \alpha \rightarrow \alpha \vdash x : (\beta \rightarrow \beta) \rightarrow (\beta \rightarrow \beta) \quad x : \forall \alpha. \alpha \rightarrow \alpha \vdash x : \beta \rightarrow \beta \\
\hline
x : \forall \alpha. \alpha \rightarrow \alpha \vdash x : \forall \alpha. \alpha \rightarrow \alpha \quad x : \forall \alpha. \alpha \rightarrow \alpha \vdash x : \forall \alpha. \alpha \rightarrow \alpha
\end{array}$$

Although  $\lambda x. xx$  is typable, the paradoxical combinator  $\Omega = (\lambda x. xx)(\lambda x. xx)$  is not, and neither is the  $Y$  combinator. This is because the language is still strongly normalizing. This means that the polymorphic  $\lambda$ -calculus is not Turing complete, that is, it cannot simulate arbitrary Turing machines.

Worse, types inference is undecidable, so the programmer must sometimes provide types.

## 5 Let-Polymorphism

We can regain decidability of type inference by placing some restrictions on the use of the type quantifier  $\forall \alpha$ . Specifically, we will only allow it at the top level; that is, we will only allow polymorphic type expressions of the form  $\forall \alpha_1 \dots \forall \alpha_n. \tau$ , where  $\tau$  is quantifier-free:

$$\begin{array}{ll}
\text{quantifier-free terms} & \tau ::= \text{int} \mid \text{bool} \mid \alpha \mid \tau_1 \rightarrow \tau_2 \\
\text{polymorphic terms} & \pi ::= \tau \mid \forall \alpha. \pi
\end{array}$$

We will also modify our rules so that it can only be introduced in the context of a `let` statement. Thus we will modify our definition of terms to include a `let` statement:

$$e ::= \dots \mid \text{let } x = e_1 \text{ in } e_2$$

and replace the generalization rule with the `let` rule

$$\frac{\Gamma \vdash d : \sigma \quad \Gamma, x : \forall \alpha_1 \dots \forall \alpha_n. \sigma \vdash e : \tau}{\Gamma \vdash \text{let } x = d \text{ in } e : \tau} \quad (\{\alpha_1, \dots, \alpha_n\} = FV(\sigma) - FV(\Gamma))$$

So type schemas are only used to type `let` expressions. For this reason, this approach is called *let-polymorphism*.

The type systems of OCaml and Haskell are based on `let`-polymorphism. We previously considered the expression `let  $x = d$  in  $e$`  to be syntactic sugar for  $(\lambda x. e) d$ , but in OCaml, the former may be typable in some cases when the latter is not:

```

# let f = fun x -> x in if (f true) then (f 3) else (f 4);;
- : int = 3
# (fun f -> if (f true) then (f 3) else (f 4)) (fun x -> x);;
Error: This expression has type int but an expression was expected of type
      bool

```

In theory, `let`-polymorphism can cause the type checker to run in exponential time, but in practice this is not a problem.

## 6 System F

In the Church-style simply-typed  $\lambda$ -calculus, we annotated binding occurrences of variables with their types. The corresponding version of the polymorphic  $\lambda$ -calculus is called *System F*. Here we explicitly abstract terms

with respect to types and explicitly instantiate by applying an abstracted term to a type. We augment the syntax with new terms and types:

$$e ::= \dots \mid \Lambda\alpha.e \mid e\tau \qquad \tau ::= b \mid \tau_1 \rightarrow \tau_2 \mid \alpha \mid \forall\alpha.\tau$$

where  $b$  are the base types (e.g., `int` and `bool`). The new terms are *type abstraction* and *type application*, respectively. Operationally, we have

$$(\Lambda\alpha.e)\tau \rightarrow e\{\tau/\alpha\}. \quad (3)$$

This just gives the rule for instantiating a type schema. Since these reductions only affect the types, they can be performed at compile time.

The typing rules for these constructs need a notion of well-formed type. We introduce a new environment  $\Delta$  that maps type variables to their *kinds*. For now, there is only one kind, namely `type`, so  $\Delta$  is a partial function with finite domain mapping type variables to  $\{\text{type}\}$ . Since the range is only a singleton, all  $\Delta$  does for now is to specify a set of types, namely  $\text{dom } \Delta$  (it will get more complicated later). As before, we use the notation  $\Delta, \alpha : \text{type}$  for the partial function  $\Delta[\text{type}/\alpha]$ . For now, we just abbreviate this by  $\Delta, \alpha$ .

The type system has two classes of judgments:

$$\Delta \vdash \tau : \text{type} \qquad \Delta; \Gamma \vdash e : \tau$$

For now, we just abbreviate the former by  $\Delta \vdash \tau$ . These judgments just determine when  $\tau$  is well-formed under the assumptions  $\Delta$ . The typing rules for this class of judgments are:

$$\Delta, \alpha \vdash \alpha \qquad \Delta \vdash b \qquad \frac{\Delta \vdash \sigma \quad \Delta \vdash \tau}{\Delta \vdash \sigma \rightarrow \tau} \qquad \frac{\Delta, \alpha \vdash \tau}{\Delta \vdash \forall\alpha.\tau}$$

Right now, all these rules do is use  $\Delta$  to keep track of free type variables. One can show that  $\Delta \vdash \tau$  iff  $FV(\tau) \subseteq \text{dom } \Delta$ .

The typing rules for the second class of judgments are:

$$\frac{\Delta \vdash \tau}{\Delta; \Gamma, x : \tau \vdash x : \tau} \qquad \frac{\Delta; \Gamma \vdash e_0 : \sigma \rightarrow \tau \quad \Delta; \Gamma \vdash e_1 : \sigma}{\Delta; \Gamma \vdash (e_0 e_1) : \tau} \qquad \frac{\Delta; \Gamma, x : \sigma \vdash e : \tau \quad \Delta \vdash \sigma}{\Delta; \Gamma \vdash (\lambda x : \sigma. e) : \sigma \rightarrow \tau}$$

$$\frac{\Delta; \Gamma \vdash e : \forall\alpha.\tau \quad \Delta \vdash \sigma}{\Delta; \Gamma \vdash (e\sigma) : \tau\{\sigma/\alpha\}} \qquad \frac{\Delta, \alpha; \Gamma \vdash e : \tau}{\Delta; \Gamma \vdash (\Lambda\alpha.e) : \forall\alpha.\tau} \quad (\alpha \notin FV(\Gamma))$$

One can show that if  $\Delta; \Gamma \vdash e : \tau$  is derivable, then  $\tau$  and all types occurring in annotations in  $e$  are well-formed. In particular,  $\vdash e : \tau$  only if  $e$  is a closed term and  $\tau$  is a closed type, and all type annotations in  $e$  are closed types.

For example, the polymorphic identity function is  $\Lambda\alpha.\lambda x : \alpha.x$ , which has polymorphic type  $\forall\alpha.\alpha \rightarrow \alpha$  according to the following proof:

$$\frac{\frac{\frac{\alpha \vdash \alpha}{\alpha; x : \alpha \vdash x : \alpha} \quad \alpha \vdash \alpha}{\alpha; \vdash (\lambda x : \alpha.x) : \alpha \rightarrow \alpha}}{\vdash (\Lambda\alpha.\lambda x : \alpha.x) : \forall\alpha.\alpha \rightarrow \alpha}$$

To apply this function to a value of a particular type, one must explicitly instantiate the type using (??):

$$((\Lambda\alpha.\lambda x : \alpha.x) \text{ int}) 3 \rightarrow (\lambda x : \text{int}.x) 3 \rightarrow 3.$$